A multiparty multi-session logic

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Abstract. Recent work on the enhancement of typing techniques for multiparty sessions with logical annotations enables, not only the validation of structural properties of the conversations and on the sorts of the messages, but also properties on the actual values exchanged. However, specification and verification of mutual effects of multiple cross-session interactions are still an open problem. We introduce a multiparty logical proof system with virtual states that enables the tractable specification and validation of fine-grained inter-session correctness properties of processes participating in several interleaved sessions. We present a sound and relatively complete static verification method.

1 Introduction

In extensively distributed computing environments, application scenarios often centre around structured conversations among multiple distributed participants. A fundamental challenge is to establish an effective specification and verification method to ensure safety in distributed software, where correctness depends on the state of individual participants and span over multiple conversations and applications. This requirement emerged from our ongoing collaboration with the Ocean Observation Initiative OOI [24], an NSF program to develop a long-term computational infrastructure for environmental ocean observation. The principals within the OOI infrastructure perform interactive activities involving distributed resources, e.g., remote instruments, off-shore sensors, data. It is important to: (1) ensure that the principals carry out each activity (session) in a way that conform a well-defined protocol, (2) express properties that span the single activities (e.g., associate each principal with a credit for resource usage, and ensure that this will always be non negative across sessions[1]).

A promising direction is the logical elaboration of types for programming languages [18]. Types offer a stable linkage between the fundamental dynamics of programs and their mathematical abstractions, serving as a highly effective basis for safety assurance. In the context of process algebras, approaches like [6, 14, 20] allow tractable[2] (e.g., with respect to model checking techniques) validation of properties such as session fidelity, progress, and error freedom. Furthermore, they enable the specification of *global* properties of multiparty interactions, yet enabling modular *local* verification of each principal. The key idea is that conversations are built as the composition of units of design called *sessions* which are specified from a global perspective (i.e., a global session type). Each global type is then *projected*, making the responsibilities of each

[1] This example is taken from the OOI Instrument Control case study and is illustrated in the Appendix of **??**.
[2] In [14, 20] verification is decidable and has linear complexity.

endpoint explicit. Validation guarantees that when each endpoint conforms to its projected specification(s), the resulting conversation conforms to the corresponding global specification(s).

These approaches require to build applications starting from a set of global types that have to be agreed upon by the principals in the network. This assumption, which poses some limitations to the flexibility with which the single local processes are modelled, is reasonable in many scenarios, provided that local processes can be built as the composition of multiple, possibly interleaved, types of sessions. However, one limitation of these approaches is that the properties that they verify are confined to the single multiparty sessions and do not treat stateful specifications incorporating mutual effects of multiple sessions run by a principal.

This paper presents a simple but powerful extension of multiparty session specifications, by enriching the assertion language studied in [6] with capability to refer to *virtual states* local to each network principal. The resulting protocol specifications are called *multiparty stateful assertions* (*MPSA*s), and model the skeletal structure of the interactions of a session, the constraints on the exchanged messages and on the branches to be followed, and the *effects* of each interaction on the virtual state. We use *invariants* to express properties, on the state of each principal, that must hold even when several sessions are executed in parallel. Principals in a network hence serve as units of verification: static validation ensures that principals behave as prescribed by *MPSA*s and their invariants are satisfied.

To see the kind of properties we are interested in, consider the following fragment of specification for the dialogue between a ticket allocation server (`S`) and its client (`C`), where the server allocates numbered tickets of increasing value to each client in consecutive, *separate* sessions:

$$\mathtt{S} \to \mathtt{C} : (y{:}\mathtt{int})\{y = \mathtt{S.x}\}\langle \mathtt{S.x{+}{+}} \rangle$$

The protocol between the server and each client is the single message-passing action where `S` sends `C` a message of type `int`. The description of this simple distributed application implies behavioural constraints of greater depth than the basic communication actions. The (sender-side) *predicate and effect* for the interaction step, $\{y = \mathtt{S.x}\}\langle \mathtt{S.x{+}{+}} \rangle$, asserts that the message y sent to each client must equal the current value of `S.x`, a state variable x allocated to the *principal* serving as `S`; and that the local effect of this message send is to increment `S.x`. In this way, `S` is specified to send incremental values across *consecutive* sessions.

The behaviour described above cannot be encoded by only using the primitives in [6]. In fact, in order to ensure inter-session properties one must discipline concurrent state updates with some mechanism of lock/unlock or atomic access/update, but lock-/unlock and atomic access/update can only be described as properties that span over multiple sessions.

Contribution We present a sound and relatively complete validation method for *MPSA*s, based on statically-verifiable proof rules. The most distinctive feature with respect to [6] is the possibility of expressing properties that span several sessions. The decidability/complexity of verification depends on the decidability/complexity of predicate evaluation in the logic that is chosen to express constraints and invariants (Proposition 10). We prove that our analysis is sound (Theorem 13) and complete (Theorem 14)

w.r.t. the semantical satisfaction relation induced by the two labelled transition systems for processes and specifications.

Synopsis In § 2 we present *MPSA*s, that is the language for (stateful) protocol specifications, and their consistency criteria (i.e., well-assertedness). In § 3 we present the calculus for networks of principals, each running a process. In § 4 we give the validation rules of networks against *MPSA*s; their properties are presented in § 5. Related work is discussed in § 6. Use cases from [24], auxiliary definitions, and full proofs can be found in the online report [5].

2 Multiparty assertions with virtual states

In the proposed framework, applications are built as the composition of units of design called *sessions*. Each type of session is specified as a *MPSA*, that is an abstract description of the interactions of the roles of a multiparty session.

The syntax of *MPSA*s is given in Figure 1. *Global assertions* ($\mathcal{G}, \mathcal{G}', \ldots$) describe a multiparty session from a global perspective; and *local assertions* ($\mathcal{L}, \mathcal{L}', \ldots$) describe it from the perspective of one role. U are types of the message contents, which can be sorts S or local assertions $\langle\mathcal{L}\rangle$ (i.e., for delegation).

$A ::= \mathsf{true} \mid \mathsf{false} \mid e_1 = e_2 \mid \neg A \mid A_1 \wedge A_2 \mid \exists x.A, \quad U ::= S \mid \langle\mathcal{L}\rangle, \quad S ::= \mathtt{bool} \mid \mathtt{int} \mid ..$

$\mathcal{G} ::=$			$\mathcal{L} ::=$		
	$\mathtt{p} \to \mathtt{q} : \{l_i(x_i : U_i)\{A_i\}\langle E_i\rangle.\mathcal{G}_i\}_{i\in I}$	**(G-int)**		$\mathtt{p}!\{l_i(x_i : U_i)\{A_i\}\langle E_i\rangle.\mathcal{L}_i\}_{i\in I}$	**(L-sel)**
$\mid$	$\mathcal{G}_1 \mid \mathcal{G}_2$	**(G-par)**	$\mid$	$\mathtt{p}?\{l_i(x_i : U_i)\{A_i\}\langle E_i\rangle.\mathcal{L}_i\}_{i\in I}$	**(L-bra)**
$\mid$	$\mu \mathsf{t}\langle y : A'\rangle(x : S)\{A\}.\mathcal{G}$	**(G-def)**	$\mid$	$\mu \mathsf{t}\langle y : A'\rangle(x : S)\{A\}.\mathcal{L}$	**(L-def)**
$\mid$	$\mathsf{t}\langle y : A'\rangle$	**(G-call)**	$\mid$	$\mathsf{t}\langle y : A'\rangle$	**(L-call)**
$\mid$	end	**(G-end)**	$\mid$	end	**(L-end)**

Fig. 1. Global and local MPSAs

For expressing constraints we use *predicates* ($A, A', \ldots$) with the syntax illustrated in Figure 1, although we may use other predicates than equality in examples. Predicates are defined on *interaction variables*, modelling the content of a message exchanged by the roles in the session, and on *state variables*, which are variables of the virtual state local to one role.

Global Assertions Interaction **(G-int)**, $\mathtt{p} \to \mathtt{q} : \{l_i(x_i : U_i)\{A_i\}\langle E_i\rangle.\mathcal{G}_i\}_{i\in I}$, models a message exchange where role $\mathtt{p}$ sends $\mathtt{q}$ one of the branch labels l_i and an interaction variable x_i, with x_i binding its occurrences in A_i, E_i, and $\mathcal{G}_i$. A_i is the predicate which needs to hold for $\mathtt{p}$ to select l_i, and which may constrain the values to be sent for x_i. Note that A_i is at the same time an assumption for the receiver $\mathtt{q}$ and a constraint for the sender $\mathtt{p}$ (i.e., if A_i is violated then the blame is on $\mathtt{p}$). E_i is the update prescribed on the virtual states of $\mathtt{p}$ and $\mathtt{q}$, modelling the persistent effects (i.e., with respect to the lifetime of the single session) of that interaction. An update is a vector of assignments of the form $\mathtt{x} := e$, where $\mathtt{x}$ is updated by the result of evaluating e in the current state.

We assume E does not contain two assignments to the same state variable, and is an atomic action. Assertion **(G-par)** is for parallel composition. The recursive definition **(G-def)** is the guarded recursion definition and defines a recursion parameter x initially set equal to a value satisfying the initialisation predicate A', with A being an invariant predicate. Global assertions are unfolded implicitly, following an equi-recursive view on types. **(G-call)** is the recursive instantiation and **(G-end)** is the termination.

Hereafter we omit true predicates, empty vectors of variables/updates, and labels of single branches.

Example 1. Consider a session with two roles, C and S. C makes an offer x to S for buying a ticket; S either accepts or refuses the offer. In the former case C spends x credits and receives a ticket, and S earns x credits. Tickets are modelled as serial numbers; they must all be increasing numbers not exceeding 1000. $\mathcal{G}_T$ below specifies this scenario:

$$\begin{aligned}\mathcal{G}_T \;&= \mathtt{C} \to \mathtt{S} : (x : \mathtt{int})\{x \geqslant 0 \wedge \mathtt{C.credit} \geqslant x\}\langle \mathtt{C.credit} := \mathtt{C.credit} - x\rangle.\\ &\quad\; \mathtt{S} \to \mathtt{C} : \{\mathsf{ok}(y : \mathtt{int})\{\mathtt{S.count} < 1000 \wedge y = \mathtt{S.count}\}\langle E_{ok}\rangle.\mathsf{end},\\ &\qquad\qquad \mathsf{ko}\langle \mathtt{C.credit} := \mathtt{C.credit} + x\rangle.\mathsf{end}\ \}\\ E_{ok} &= \mathtt{S.credit} := \mathtt{S.credit} + x, \mathtt{S.count} := \mathtt{S.count} + 1\end{aligned}$$

C has state variable `credit`, and S has state variables `credit` and `count` (a counter for serial numbers). The first interaction requires that the offer x does not exceed C's credit, and decrements the credit by x. S selects one of the two branches by either label ok or ko. The former branch can be selected only if $\mathtt{S.count} < 1000$.

We denote with $var(\mathcal{G})$ the set of (interaction/state) variables and recursion parameters in $\mathcal{G}$, and with $var(A)$ the free variables of A (same for e). The set of variables that $\mathtt{p} \in \mathcal{G}$ *knows*, written $var(\mathcal{G}) \upharpoonright \mathtt{p}$, consists of: (i) the state variables of the form $\mathtt{p}.x$ for some x, (ii) the interaction variables sent or received by p in $\mathcal{G}$, and (iii) the parameters of the recursive definitions $\mu \mathsf{t}\langle y : A'\rangle(x : S)\{A\}.\mathcal{G}'$ in $\mathcal{G}$ such that p knows all the free variables in initialisation A', and all free variables in A'' for all $\mathsf{t}\langle y : A''\rangle$ in $\mathcal{G}'$ (we assume each recursion parameter known by exactly two participants).

Well-assertedness Our theory relies on two consistency principles: *history-sensitivity* and *temporal-satisfiability*. These principles were first introduced in [6]; we discuss them here as their adaptation to our stateful scenario requires non-trivial extensions.

By history-sensitivity each role must have enough information to fulfil the specified obligations, namely it requires that: (1) each role p knows all free variables in the predicates that p must guarantee, and (2) each role has enough information to perform the prescribed updates, that is (i) when to make an update, and (ii) which values to assign.

Definition 2 (History-sensitivity). $\mathcal{G}$ is history-sensitive if for each interaction, of the form $\mathtt{p} \to \mathtt{q} : \{l_i(x_i : U_i)\{A_i\}\langle E_i\rangle.\mathcal{G}_i\}_{i\in I}$, occurring in $\mathcal{G}$, for all $i \in I$:

1. $var(\mathcal{G}) \upharpoonright \mathtt{p} \supseteq var(A_i)$ (i.e., p knows all variables in $var(A_i)$),
2. for all $\mathtt{r}.x := e$ in E_i: (i) either $\mathtt{r} = \mathtt{p}$ or $\mathtt{r} = \mathtt{q}$, and (ii) $var(\mathcal{G}) \upharpoonright \mathtt{r} \supseteq var(e)$.

A checker for history-sensitivity can be found in Appendix D.1. Consider the assertions:

$$\begin{aligned}\mathcal{G} \;&= \mathtt{p} \to \mathtt{q} : (x{:}\mathtt{int}).\ \mathtt{q} \to \mathtt{r} : (y{:}\mathtt{int}).\ \mathtt{r} \to \mathtt{s} : (z{:}\mathtt{int})\{z > x\}\\ \mathcal{G}' &= \mathtt{p} \to \mathtt{q} : (y{:}\mathtt{int}).\ \mathtt{q} \to \mathtt{r} : \{\mathsf{ok}(w{:}\mathtt{int})\langle \mathtt{r.x}_1 := y, \mathtt{p.x}_2 := y\rangle, \mathsf{ko}\}\end{aligned}$$

$\mathcal{G}$ violates (1) because $\mathtt{r}$ has to send a value for z that is greater than x without knowing x. $\mathcal{G}'$ violates both clauses of (2): (i) because $\mathtt{p}$ must update $\mathtt{x}_2$ not knowing whether and when the update should be done, and (ii) because in the second interaction $\mathtt{r}$ has to update $\mathtt{x}_1$ with y without knowing y.[3]

By temporal-satisfiability, for each participant $\mathtt{p} \in \mathcal{G}$, whenever it is $\mathtt{p}$'s turn to send a value, $\mathtt{p}$ can find at least one selection branch and one value which satisfies the specified constraint. Temporal satisfiability is defined (and checked) using a function $\mathtt{ts}(\mathcal{G}, A)$ which returns true only if $\mathcal{G}$ always allows a path of interactions going through $\mathcal{G}$ in *any possible state*. Considering all possible states makes the specification robust with respect to arbitrary interactions the same principal may be engaged in through other sessions. Predicate A is incrementally built as a conjunction of the predicates that appear in $\mathcal{G}$ in all the recursive invocations and models the current set of assumptions.

Definition 3 (Temporal-satisfiability). Let $\mathcal{G}$ be a global specification, and A a predicate. $\mathtt{ts}(\mathcal{G}, A)$ is given by:

1. $\mathtt{ts}(\mathtt{p}\to\mathtt{q} : \{l_i(x_i : U_i)\{A_i\}\langle E_i\rangle.\mathcal{G}_i\}_{i\in I}, A) = \begin{cases} \bigwedge_{i\in I} \mathtt{ts}(\mathcal{G}_i, A \wedge \underline{A_i}) & \text{if } A \supset \bigvee_{i\in I} \exists x_i.A_i \\ \mathsf{false} & \text{otherwise} \end{cases}$
2. $\mathtt{ts}(\mathcal{G}_1 \mid \mathcal{G}_2, A) = \mathtt{ts}(\mathcal{G}_1, A) \wedge \mathtt{ts}(\mathcal{G}_2, A)$
3. $\mathtt{ts}(\mu \mathtt{t}\langle e\rangle(x\!:\!S)\{A'\}.\mathcal{G}', A) = \begin{cases} \mathtt{ts}(\mathcal{G}', A \wedge A') & \text{if } A \supset (A'[e/x]) \\ \mathsf{false} & \text{otherwise} \end{cases}$
4. $\mathtt{ts}(t_{A'(x)}\langle e\rangle, A) = \begin{cases} \mathsf{true} & \text{if } A \supset A'[e/x] \\ \mathsf{false} & \text{otherwise} \end{cases}$
5. $\mathtt{ts}(\mathsf{end}, A) = \mathsf{true}$

$\mathcal{G}$ satisfies temporal satisfiability if $\mathtt{ts}(\mathcal{G}, \mathsf{true}) = \mathsf{true}$.[4]

In (1) the first condition for "if" demands that there exists at least one branch for which it is possible to find a value for x_i that satisfies the current predicate A_i. The function is called recursively extending the set of preconditions A with with the closure $\underline{A_i}$ of predicate A_i (see Remark 4 below). (2) demands both parts of the composition are satisfiable. (3) and (4) check recursion, the latter relying on the annotation of recursive calls with the invariants of the corresponding recursive definitions.

Remark 4. The closure of a predicate A in $\mathcal{G}$, written $\underline{A}$, is the predicate obtained by closing with existential quantifiers the free state variables of $\mathcal{G}$ in A. Whereas the values of interaction variables in a session do not change after they are introduced[5], state variables can be updated a number of times in a number of different ways. Hence

[3] [7] proposes algorithms to amend assertions that violate history-sensitivity and temporal-satisfiability as in [6]. No such algorithms have yet been investigated for the definitions introduced in this paper. Although relevant, the issue of amending inconsistent assertions is out of the scope of the current work.

[4] This property can be relaxed by starting from a stronger precondition A as long as A is then implied by the principal invariants (which are defined in § 4).

[5] Actually, interaction variables in a recursion body are reused at each iteration. However, their values are due to follow the same constraints at each iteration.

a predicate on state variables may be true at a certain time, and become false at a later time. Hereafter we use $\underline{A}$ when we want to to 'keep' only the persistent assumptions (those on interaction variables) of A.

The following global specification violates temporal satisfiability

$$\mathtt{p} \to \mathtt{q} : (x : \mathtt{int})\{x > 0\}.\mathtt{q} \to \mathtt{p} : (y : \mathtt{int})\{y = x \wedge y > 100\}$$

In fact, in the first interaction p is allowed to choose any positive value for x, for instance 10. In this case, q cannot find any value for y such that $y = 10 \wedge y > 100$.

Proposition 5. *Given a global assertion $\mathcal{G}$, let m be the size of the syntactic tree of $\mathcal{G}$, n be the maximum number of variables occurring in each predicate in $\mathcal{G}$, and $eval(A)$ be the complexity of predicate evaluation (if decidable). History-sensitivity can be checked in $O(m \times n)$. Temporal-satisfiability is decidable if predicate evaluation is decidable and, if decidable, it can be checked in $O(m) \times eval(A)$.*

Hereafter, we assume assertions to be *well-asserted*.

Local Assertions Each local assertion $\mathcal{L}$ refers to a specific role. Syntax is given in the right part of Figure 1. Selection **(L-sel)** $\mathtt{p}!\{l_i(x_i : U_i)\{A_i\}\langle E_i\rangle.\mathcal{L}_i\}_{i\in I}$ models an interaction where the role sends p a branch label l_i and a message x_i. A_i and E_i are the predicate and update respectively. **(L-bra)** is the dual branching. The others are as in the global assertions, except that a local assertion cannot be multi-threaded.

Given a global assertion $\mathcal{G}$, we can automatically derive the local assertions for each role $\mathtt{p} \in \mathcal{G}$ by *projection*. The projection rules rely on a few auxiliary definitions: projection of a predicate, and projection of an update. The *projection of a predicate* A on p in $\mathcal{G}$, written $A \upharpoonright \mathtt{p}$, is defined as $\exists\tilde{x}.A$ where $\tilde{x} = var(A)\backslash(var(\mathcal{G}) \upharpoonright \mathtt{p})$ (i.e., the existential closure of the variables that p does not know). The *projection of an update* E on p in $\mathcal{G}$, written $E \upharpoonright \mathtt{p}$ is the update E' containing only the assignments $\mathtt{p}_j.\mathtt{x}_i := e_j$ such that $\mathtt{p}_j = \mathtt{p}$.

The projection rules for global assertions are as in [6], except that updates are now considered; their detailed presentation is not necessary to understand the results in this paper, hence we only give an illustration through Example 6. Henceforth, in $\mathcal{G} \upharpoonright \mathtt{p}$ we shall omit the $\mathtt{p}.\_$ prefix when referring to p's state variables.

Example 6. $\mathcal{L}_\mathtt{C}$ (resp. $\mathcal{L}_\mathtt{S}$) is the projection of $\mathcal{G}_T$ from Example 1 on C (resp. S).

$$\begin{aligned}
\mathcal{L}_\mathtt{C} &= \mathtt{S}\,!\,(x : \mathtt{int})\{x \geqslant 0 \wedge \mathtt{credit} \geqslant \mathtt{x}\}\langle \mathtt{credit} := \mathtt{credit} - x\rangle.\mathcal{L}'_\mathtt{C}\\
\mathcal{L}'_\mathtt{C} &= \mathtt{S}\,?\{\mathsf{ok}(y : \mathtt{int})\{\exists \mathtt{S.count}.\mathtt{S.count} < 1000 \wedge y = \mathtt{S.count}\}.\mathsf{end},\\
&\qquad \mathsf{ko}\langle \mathtt{credit} := \mathtt{credit} + x\rangle.\mathsf{end}\}\\
\mathcal{L}_\mathtt{S} &= \mathtt{C}\,?(x : \mathtt{int})\{\exists \mathtt{C.credit}.x \geqslant 0 \wedge \mathtt{C.credit} \geqslant x\}.\mathcal{L}'_\mathtt{S}\\
\mathcal{L}'_\mathtt{S} &= \mathtt{C}\,!\,\{\mathsf{ok}(y : \mathtt{int})\{\mathtt{count} < 1000 \wedge y = \mathtt{count}\}\\
&\qquad\qquad \langle \mathtt{credit} := \mathtt{credit} + x, \mathtt{count} := \mathtt{count} + 1\rangle.\mathsf{end},\\
&\qquad \mathsf{ko}.\mathsf{end}\}
\end{aligned}$$

The projection of the first interaction of $\mathcal{G}_T$ on sender C (resp. receiver S) is a send/select (resp. a receive/branch). The predicates/updates of the projections on a role are the projections of the predicates/updates on that role.[6] The continuation is projected similarly,

[6] Note that by well-assertedness (clause 1) the projection of a predicate on the sender of an interaction is always the predicate itself.

proceeding point-wise for each branch. Sometimes the projected predicate includes information about constraints of interactions between third parties (without however revealing the actual values exchanged by the third parties), e.g., $\exists \mathtt{S.count}.\mathtt{S.count} < 1000 \wedge y = \mathtt{S.count}$ provides $\mathtt{C}$ with precondition $y < 1000$.

Well-assertedness is easily extended to local assertions.

3 Multiparty networks with local states

We consider networks of interactional entities called *principals* linked by a common global transport, modelled as queues. Each principal runs a *located process*, that is a process with multiparty session primitives [2, 20] (to enable rigorous representation of conversation structures) and with a *local state*.

Syntax The syntax of networks and processes is given in Figure 2 and is a refined version of the multiparty session π-calculus from [2, 11] with local states. A local state σ maps a signature $[\tilde{\mathtt{x}} : \tilde{S}]$ of typed pairwise disjoint state variables $\tilde{\mathtt{x}}$ to their sorts. We use the injective function $\mathtt{id}(\sigma)$ to map each local state to an identifier.

A network can be an empty network $\varnothing$, a located process $[P]\sigma$, a parallel composition of networks $N_1 \mid N_2$, a new session name $(\nu s)N$ which binds s in N, or a queue $s : h$ where h are messages in transit through session channel s. A network is *initial* if it has no new session names and queues, otherwise it is *runtime*. We denote the free session channels in N with $\mathsf{fn}(N)$, similarly for P with $\mathsf{fn}([P]\sigma) = \mathsf{fn}(P)$.

$$
\begin{array}{lll}
\text{(network)} & N ::= \varnothing \mid [P]\sigma \mid N_1 | N_2 \mid (\nu s)N \mid s : h & \\
\text{(state/queue/value)} & \sigma ::= [\tilde{\mathtt{x}} : \tilde{S}] \mapsto \tilde{S} \quad h ::= \varnothing \mid (\mathtt{p}, \mathtt{q}, l\langle v \rangle) \cdot h \quad v ::= \mathtt{n} \mid s[\mathtt{p}] & \\
\text{(process)} & P ::= \overline{a}[\mathtt{n}](y).P & \textbf{(P-req)} \\
 & \mid\ a[\mathtt{i}](y).P & \textbf{(P-acc)} \\
 & \mid\ k[\mathtt{p}, \mathtt{q}]!\{e_i \mapsto l_i\langle e_i'\rangle(x_i)\langle E_i\rangle; P_i\}_{i\in I} & \textbf{(P-sel)} \\
 & \mid\ k[\mathtt{p}, \mathtt{q}]?\{l_i(x_i)\langle E_i\rangle.P_i\}_{i\in I} & \textbf{(P-bra)} \\
 & \mid\ P \mid Q & \textbf{(P-par)} \\
 & \mid\ (\mu X(x).P)\langle e\rangle & \textbf{(P-def)} \\
 & \mid\ X\langle e\rangle & \textbf{(P-call)} \\
 & \mid\ \mathbf{0} & \textbf{(P-idle)} \\
\text{(channel/update/exp)} & k ::= y \mid s \quad E ::= \varnothing \mid E; \mathtt{x} := e \quad e ::= v \mid e\ \mathtt{op}\ e &
\end{array}
$$

$x, y, \ldots$ interaction variables $\quad \mathtt{x}, \mathtt{y}, \ldots$ state variables $\quad X, Y, \ldots$ process variables
$a, b, \ldots$ shared name $\quad s, s', \ldots$ session name $\quad \mathtt{n}, \mathtt{n}', \ldots$ constants

Fig. 2. Syntax of networks and processes

Session request **(P-req)** multicasts a request to each session accept process **(P-acc)**, e.g., $a[\mathtt{i}](y).P$ with $i \in \{2, .., n\}$, by synchronisation through a shared name a and continuing as P. **(P-sel)** is Dijkstra's guarded command [17] and **(P-bra)** is the branching process; they represent communications through an established session k. **(P-sel)** acts as role $\mathtt{p}$ in session k and sends role $\mathtt{q}$ one of the labels l_i. The choice of the label is determined by boolean expressions e_i, assuming $\vee_{i\in I} e_i = \mathsf{true}$ and $i \neq j$ implies

$e_i \wedge e_j = \mathsf{false}$. Each label l_i is sent with the corresponding expression e'_i which specifies the value for x_i, assuming e'_i and x_i have the same type.[7] **(P-bra)** plays role $\mathtt{q}$ in session k and is ready to receive from $\mathtt{p}$ one of the labels l_i and a value for the corresponding x_i, then behaves as P_i after instantiating x_i with the received value. In guarded command (resp. branching), the local state of the sender (resp. receiver) is updated according to update E_i; in both processes each x_i binds its occurrences in P_i and E_i.

Example 7. Processes $P_{\mathtt{S}}$ and $P_{\mathtt{C}}$ implement $\mathcal{L}_{\mathtt{S}}$ and $\mathcal{L}_{\mathtt{C}}$, respectively, from Example 6.

$$P_{\mathtt{S}} = a[\mathtt{2}](z).z[\mathtt{C},\mathtt{S}]?(x); P'_{\mathtt{S}} \qquad E_{ok} = \mathtt{count} := \mathtt{count} + 1, \mathtt{credit} := \mathtt{credit} + x$$
$$P'_{\mathtt{S}} = z[\mathtt{S},\mathtt{C}]!\{\{\mathtt{count} < 1000 \wedge x \geqslant 10\} \mapsto \mathsf{ok}\langle \mathtt{count}\rangle(y)\langle E_{ok}\rangle.\mathbf{0},\ \{\mathtt{count} \geqslant 1000 \vee x < 10\} \mapsto \mathsf{ko}.\mathbf{0}\}$$
$$P_{\mathtt{C}} = \overline{a}[\mathtt{2}](w).w[\mathtt{C},\mathtt{S}]!\langle 8\rangle\langle x\rangle\langle \mathtt{credit} := \mathtt{credit} - x\rangle; P'_{\mathtt{C}}$$
$$P'_{\mathtt{C}} = w[\mathtt{S},\mathtt{C}]?\{\mathsf{ok}(y).\mathbf{0},\ \mathsf{ko}\langle \mathtt{credit} := \mathtt{credit} + x\rangle.\mathbf{0}\}$$

We let $\mathtt{C} = \mathtt{1}$ and $\mathtt{S} = \mathtt{2}$. $P_{\mathtt{S}}$ accepts a request to participate to a session specified by $\mathcal{G}_T$ (assuming a has type $\mathcal{G}_T$) on channel z as role $\mathtt{2}$. In the established session z, the principal receives an offer x from the co-party. It follows a guarded command with two cases; if $\mathtt{count}$ has not reached its maximum value for serial numbers and the offer is greater than 10 then the first branch (ok) is taken and $\mathtt{count}$ is sent as y, otherwise the second branch (ko) is taken. Dually, $P_{\mathtt{C}}$ sends a request to participate to one instance of session $\mathcal{G}_T$ as the role $\mathtt{1}$. A principal may repeatedly execute a process using recursion, or run concurrent instances of the same type of session (e.g., $[P_{\mathtt{S}} \mid P_{\mathtt{S}}]\sigma$) or different types of session (e.g., $[P_{\mathtt{S}} \mid P_{\mathtt{C}}]\sigma$) as discussed in Example 9.

Operational semantics The LTS is generated from the rules in Figure 3 using the following labels: $\ell ::= \overline{a}[\mathtt{n}]\langle s\rangle \mid a[\mathtt{i}]\langle s\rangle \mid s[\mathtt{p},\mathtt{q}]!l\langle v\rangle \mid s[\mathtt{p},\mathtt{q}]?l\langle v\rangle \mid \tau$. We denote with $\sigma\ \mathtt{after}\ E$ the state σ after the update E. We write $\sigma \models e \downarrow v$ for a closed expression e when it evaluates to v in σ.
The first and second rule are for requesting and accepting a session initialisation. The guarded command checks if condition e_j is satisfied in the current state σ, and sends a message consisting of one of the labels l_j and an expression e'_j (which is evaluated to a value v in state σ), updates σ according to E_j, and behaves as $P[v/x_j]$. Branching is symmetric. The synchronous session initialisation creates a new queue. We omit the standard context/structural equivalence rules.

4 Proof system for multiparty session logic with virtual states

In this section we outline how to obtain the syntactic validation of networks, written $\Gamma \vdash N \rhd \Sigma$, assuming processes typable, following [6]. The proof rules rely on the following environments:

$$\Gamma ::= \varnothing \mid \Gamma, a : \mathcal{G} \mid \Gamma, X : (x : S)\mathcal{L}_1 @ \mathtt{p}_1, \ldots, \mathcal{L}_n @ \mathtt{p}_n, \qquad \Delta ::= \varnothing \mid \Delta, s[\mathtt{p}] : \mathcal{L},$$
$$\Sigma ::= \varnothing \mid \Sigma, [\Delta]\sigma$$

[7] Guarded command can be implemented using selection, if-then-else and lock-unlock. Although our theory is applicable to these primitives, we choose to make these low-level steps atomic for minimising the syntax.

$$[\overline{a}[\mathtt{n}](y).P]\sigma \xrightarrow{\overline{a}[\mathtt{n}]\langle s\rangle} [P[s/y]]\sigma \qquad [a[\mathtt{i}](y).P]\sigma \xrightarrow{a[i]\langle s\rangle} [P[s/y]]\sigma \quad (s \notin \mathsf{fn}(P))$$

$$[s[\mathtt{p},\mathtt{q}]!\{e_i \mapsto l_i\langle e_i'\rangle(x_i)\langle E_i\rangle; P_i\}_{i\in I}]\sigma \xrightarrow{s[\mathtt{p},\mathtt{q}]!l_j\langle v\rangle} [P[v/x_j]]\sigma'$$
$$(j \in I \quad \sigma \models e_j' \downarrow v \quad \sigma \models e_j \quad \sigma' = \sigma \,\mathtt{after}\, E_j[v/x_j])$$

$$[s[\mathtt{p},\mathtt{q}]?\{l_i(x_i)\langle E_i\rangle.P_i\}_{i\in I}]\sigma \xrightarrow{s[\mathtt{p},\mathtt{q}]?l_j\langle v\rangle} [P_j[v/x_j]]\sigma' \qquad (j \in I \quad \sigma' = \sigma \,\mathtt{after}\, E_j[v/x_j])$$

$$\frac{[P_1]\sigma_1 \xrightarrow{\overline{a}[\mathtt{n}]\langle s\rangle} [P_1']\sigma_1 \quad [P_i]\sigma_i \xrightarrow{a[i]\langle s\rangle} [P_i']\sigma_i \quad (2 \leqslant i \leqslant n)}{[P_1]\sigma_1 \mid \cdots \mid [P_n]\sigma_n \xrightarrow{\tau} (\nu s)(s : \varnothing \mid [P_1']\sigma_1 \mid \cdots \mid [P_n']\sigma_n)}$$

$$\frac{[P]\sigma \xrightarrow{s[\mathtt{p},\mathtt{q}]!l_j\langle v\rangle} [P']\sigma'}{[P]\sigma \mid s : h \xrightarrow{\tau} [P']\sigma' \mid s : h \cdot (\mathtt{p},\mathtt{q}, l_j\langle v\rangle)} \qquad \frac{[P]\sigma \xrightarrow{s[\mathtt{p},\mathtt{q}]?l_j\langle v\rangle} [P']\sigma'}{[P]\sigma \mid s : (\mathtt{p},\mathtt{q}, l_j\langle v\rangle) \cdot h \xrightarrow{\tau} [P']\sigma' \mid s : h}$$

Fig. 3. Labelled transition for networks

Γ maps shared names to global assertions and process variables to their parameters. If $\Gamma \vdash a : \mathcal{G}$ then a session specified by $\mathcal{G}$ can be initiated by processes (via session request or accept) using a. By the standard kinding rules, we check if the same free variable appears in different global types in Γ, then they have the same sort. The mapping of process variables is for the validation of recursive assertions. Δ maps session channels/roles to local assertions. If $\Delta \vdash s[\mathtt{p}] : \mathcal{L}$ then a session is active (i.e., it has been initialized) on channel s for role $\mathtt{p}$; $\mathcal{L}$ specifies the (part of the) session that has still to be executed. Σ is the specification of a network; each $[\Delta]_\sigma$ is the specification of a located process with the respective virtual state.

We also use an *assertion environment* $\mathcal{C}$, which is incrementally built by conjunction of the predicates and boolean expressions (i.e., the conditions of a guarded commands) occurring in the processes being validated, and models their assumptions. Hereafter, given a predicate A and an update E, we define $A \,\mathtt{after}\, E$ to be the predicate obtained by substituting, for each assignment $x := e$ in E, each occurrence of x in A with e.

Modelling cross-session properties: the principal invariant Given a located process $[P]\sigma$ in a network, we want to allow the architect to model stable properties (i.e., invariant) over the variables in σ on across multiple sessions. We call these properties *principal invariant* of $[P]\sigma$, that is a predicate (following the syntax for A in Figure 1) over the state variables of σ. Hereafter we assume there exists a function $\mathcal{I}(\sigma)$ that given a local state σ returns the principal invariant for σ. Principal invariants depend from the application domain, and the architect should define them prior to the verification.

Example 8. Consider a located process $[P_\mathtt{c} \mid P_\mathtt{s}]\sigma_p$ with $P_\mathtt{c}$ and $P_\mathtt{s}$ from Example 7. Assume we want to require that the credit is always non-negative (i.e., the principal does not contracts debts) and that the counter does not exceed the maximum number of tickets which is 1000. We can enforce these constraints by setting the principal invariant $\mathcal{I}(\sigma_p)$ to be $\mathtt{credit} \geqslant 0 \wedge 0 \leqslant \mathtt{count} \leqslant 1000$.

Proof rules Figure 4 illustrates the proof rules for initial networks and processes.

$$\frac{\mathtt{id}(\sigma_p) = \mathtt{id}(\sigma_a) \quad \sigma_p, \sigma_a \models \mathcal{I}(\sigma_p) \quad \mathcal{I}(\sigma_p) \wedge \mathcal{C}; \Gamma \vdash P \rhd \Delta}{\mathcal{C}; \Gamma \vdash [P]\sigma_p \rhd [\Delta]\sigma_a} \quad \lfloor \text{N1} \rfloor$$

$$\frac{(\Gamma', \Delta', \sigma') \ni (\Gamma, \Delta, \sigma) \quad \mathcal{C} \supset \mathcal{C}' \quad \mathcal{C}'; \Gamma' \vdash N \rhd [\Delta']\sigma'}{\mathcal{C}; \Gamma \vdash N \rhd [\Delta]\sigma} \quad \lfloor \text{N2} \rfloor$$

$$\frac{-}{\mathcal{C}; \Gamma \vdash \varnothing \rhd \varnothing} \qquad \frac{\mathcal{C}; \Gamma \vdash N \rhd \Sigma \qquad \mathcal{C}; \Gamma \vdash N' \rhd \Sigma'}{\mathcal{C}; \Gamma \vdash N \mid N' \rhd \Sigma, \Sigma'} \quad \lfloor \text{N3/N4} \rfloor$$

$$\frac{\mathcal{C}; \Gamma, a : \mathcal{G} \vdash P \rhd y[\mathtt{i}] : \mathcal{G} \upharpoonright \mathtt{i}, \Delta}{\mathcal{C}; \Gamma, a : \mathcal{G} \vdash a[\mathtt{i}](y).P \rhd \Delta} \quad \lfloor \text{MACC} \rfloor$$

$$\frac{\forall i \in I, \qquad \mathcal{C} \wedge A_i; \Gamma \vdash P_i \rhd \Delta, k[\mathtt{q}] : \mathcal{L}_i \qquad \mathcal{C} \wedge A_i \supset (\mathcal{C} \,\mathtt{after}\, E_i)}{\mathcal{C}; \Gamma \vdash k[\mathtt{p}, \mathtt{q}]?\{l_i(x_i)\langle E_i\rangle.P_i\}_{i \in I} \rhd \Delta, k[\mathtt{q}] : \mathtt{p}?\{l_i(x_i : U_i)\{A_i\}\langle E_i\rangle.\mathcal{L}_i\}_{i \in I}} \quad \lfloor \text{BCH} \rfloor$$

$$\frac{\begin{array}{c}\forall i \in I \exists j \in J, \qquad l_i = l_j \qquad \mathcal{C} \wedge \underline{e_i}; \Gamma \vdash P_i[e_i'/x_i] \rhd \Delta, k[\mathtt{p}] : \mathcal{L}_j[e_i'/x_j] \\ \mathcal{C} \wedge e_i \supset (A_j \wedge (E_i = E_j) \wedge (\mathcal{C} \,\mathtt{after}\, E_j))[e_i'/x_i]\end{array}}{\mathcal{C}; \Gamma \vdash k[\mathtt{p}, \mathtt{q}]!\{e_i \mapsto l_i\langle e_i'\rangle(x_i)\langle E_i\rangle; P_i\}_{i \in I} \rhd \Delta, k[\mathtt{p}] : \mathtt{q}!\{l_j(x_j : U_j)\{A_j\}\langle E_j\rangle.\mathcal{L}_j\}_{j \in J}} \quad \lfloor \text{SEL} \rfloor$$

$$\frac{\mathcal{C}; \Gamma \vdash P_1 \rhd \Delta_1 \quad \mathcal{C}; \Gamma \vdash P_2 \rhd \Delta_2}{\mathcal{C}; \Gamma \vdash P_1 \mid P_2 \rhd \Delta_1, \Delta_2} \qquad \frac{\Delta \text{ end only}}{\mathcal{C}; \Gamma \vdash \mathbf{0} \rhd \Delta} \quad \lfloor \text{PAR/END} \rfloor$$

$$\frac{\mathcal{L}_1[e/x], \ldots, \mathcal{L}_n[e/x] \text{ well-asserted}}{\mathcal{C}; \Gamma, X : (x)\mathcal{L}_1 @ \mathtt{p}_1, \ldots, \mathcal{L}_n @ \mathtt{p}_n \vdash X\langle e\rangle \rhd s[\mathtt{p}_1] : \mathcal{L}_1[e/x], \ldots, s[\mathtt{p}_n] : \mathcal{L}_n[e/x]} \quad \lfloor \text{VAR} \rfloor$$

$$\frac{\mathcal{C}; \Gamma, X : (x)\mathcal{L}_1 @ \mathtt{p}_1, \ldots, \mathcal{L}_n @ \mathtt{p}_n \vdash P \rhd s[\mathtt{p}_1] : \mathcal{L}_1, \ldots, s[\mathtt{p}_n] : \mathcal{L}_n}{\mathcal{C}; \Gamma \vdash (\mu X(x).P)\langle e\rangle \rhd s[\mathtt{p}_1] : \mathcal{L}_1[e/x], \ldots, s[\mathtt{p}_n] : \mathcal{L}_n[e/x]} \quad \lfloor \text{REC} \rfloor$$

Fig. 4. Proof rules for networks (top) and proof rules for processes (bottom)

$\lfloor$N1$\rfloor$ decomposes the validation of a network into the validations of each located process against its corresponding specification Δ. The correspondence between principal and specification in checked by the clause $\mathtt{id}(\sigma_p) = \mathtt{id}(\sigma_a)$. Furthermore, local and virtual states must satisfy the principal invariant $\mathcal{I}(\sigma_p)$. P is then validated in the assertion environment extended (i.e., in conjunction with) the principal invariant.

$\lfloor$N2$\rfloor$ is the rule for refinement. This rule is useful to validate processes even if they do not match exactly a given assertion as long as they implement a behaviour that is 'more refined' than the one prescribed. Refinement is also necessary to prove completeness of these rules (Theorem 14). We use the following refinement relation between specifications: $(\Gamma', \Delta', \sigma') \ni (\Gamma, \Delta, \sigma)$ if $(\Gamma', \Delta', \sigma')$ specifies a more refined behaviour than (Γ, Δ, σ), in that it poses more restrictions on the output actions and poses less restrictions on the input actions. $\lfloor$N2$\rfloor$ allows to refine the assertion environment $\mathcal{C}$ by considering, in the premise, a weaker set of assumptions $\mathcal{C}'$.

$\lfloor$N3$\rfloor$ is for empty networks and $\lfloor$N4$\rfloor$ is for decomposing the validation of networks.

$\lfloor$MACC$\rfloor$ validates a session accept on a shared channel a as role $\mathtt{i}$ provided that a is in the domain of Γ, and that the continuation P is validated against the specification Δ extended with the new session $y[\mathtt{i}]$. In the (now active) session $y[\mathtt{i}]$, P must behave as $\Gamma(a)$ projected on role $\mathtt{i}$. The rule for session request is similar hence omitted.

$\lfloor$BCH$\rfloor$ validates the branching process. Δ must include an active session $k[\mathtt{q}]$ on session channel k for the receiver role $\mathtt{q}$. In the premise, the continuation for each branch i is required to be still valid in the assertion environment extended with A_i. In

the second clause of the premise, for each branch i the update E_i must not invalidate $\mathcal{C}$; this ensures that the update does not invalidate the principal invariant. The invariant is not mentioned explicitly (to keep the proof rules concise), but it is implied by $\mathcal{C}$. In fact, $\mathcal{C}$ is the conjunction of (1) the principal invariant (by ⌊N1⌋), (2) possibly some interaction predicates (by ⌊BCH⌋), and (3) possibly some boolean expressions (by ⌊SEL⌋). Since predicates (2), (3) and A_i do not contain free state variables[8], then E_i can only invalidate the principal invariant (1); on the other hand (2), (3) and A_i are necessary premises (i.e., $\mathcal{C} \wedge A_i$ before the implication) as they may constrain interaction variables used by E_i.

In ⌊SEL⌋ each branch i of the process must correspond to a branch j of the specification ($l_i = l_j$). The continuation must be validated in assertion environment $\mathcal{C}$ extended with the closure $\underline{e_i}$ of the condition of the branch e_i. The closure of boolean expression e_i is defined as the closure for predicates (see Remark 4). The clause at the bottom of the premise requires that, under the assumption $\mathcal{C} \wedge e_i$: (1) expression e_i' satisfies A_j, (2) assertion and process have the same effects/updates on the states, (3) update E_j does not invalidate the principal invariant. [9]

⌊PAR⌋ is similar to ⌊N2⌋ but for parallel processes. ⌊END⌋ validates the idle process provided that each active session in the specification Δ is of the form $y[\mathtt{p}] : \mathsf{end}$.

⌊VAR⌋ validates recursive call given that the active sessions in Δ correspond to the roles and local assertions associated to process variable X in Γ and that each $\mathcal{L}_i$ is still well-asserted when the recursion parameter is substituted with e. ⌊REC⌋ is the standard rule for recursion definition. The validation of recursive processes is handled in a similar way to [6]; it uses a refinement rule for processes, similar to ⌊N2⌋ and omitted for simplicity, and the fact that assertions are refined by their unfolding. See [5] for more details.

Example 9. Consider the located process $[P_{\mathtt{S}} \mid P_{\mathtt{C}}]\sigma_p$ from Example 8 that executes two parallel threads: one selling a ticket and the other one buying another kind of ticket from another principal (the other principal is not modelled here). We show the validation of $\mathsf{true}; \Gamma \vdash [P_{\mathtt{S}} \mid P_{\mathtt{C}}]\sigma_p \rhd [\varnothing]\sigma_a$ proceeding top-down using the rules in Figure 4.

The global specification $[\varnothing]\sigma_a$ is initially empty since there are no active sessions. The active sessions will be added upon session request/accept by $P_{\mathtt{S}}$ and $P_{\mathtt{C}}$. We assume $\sigma_p = \sigma_a = \{\mathtt{count} : \mathtt{int}, \mathtt{credit} : \mathtt{int}\} \mapsto \{10, 500\}$ and initially $\mathcal{C} = \mathsf{true}$.

We first apply ⌊N1⌋ with $\mathcal{I}(\sigma_p) = \mathtt{credit} \geqslant 0 \wedge 0 \leqslant \mathtt{count} \leqslant 1000$ from Example 8. For readability we will write $\mathcal{I}$ instead of $\mathcal{I}(\sigma_p)$ in this example. Note that $\mathcal{I}$ is satisfied by the local and virtual state. Next we apply rule ⌊PAR⌋ that decomposes the derivation of two threads for $P_{\mathtt{S}}$ and $P_{\mathtt{C}}$. We omit the illustration of the latter thread.

Below we illustrate the application of rule ⌊MACC⌋ and ⌊BCH⌋ to the former thread:

$$\dfrac{\dfrac{\begin{array}{c}\mathcal{I} \wedge \{\exists \mathtt{C}.\mathtt{credit}.\mathtt{C}.\mathtt{credit} \geqslant x\}; \Gamma \vdash P_{\mathtt{S}}' \rhd z[\mathtt{S}] : \mathcal{L}_{\mathtt{S}}' \\ \mathcal{I} \wedge \{\exists \mathtt{C}.\mathtt{credit}.\mathtt{C}.\mathtt{credit} \geqslant x\} \supset (\mathcal{I}\ \mathtt{after}\ \varnothing)\end{array}}{\mathcal{I}; \Gamma \vdash z[\mathtt{C},\mathtt{S}]?(x).P_{\mathtt{S}}' \rhd z[\mathtt{S}] : \mathtt{C}?(x : \mathtt{Nat})\{\exists \mathtt{C}.\mathtt{credit}.\mathtt{C}.\mathtt{credit} \geqslant x\}.\mathcal{L}_{\mathtt{S}}'}\ \lfloor \mathrm{BCH} \rfloor}{\mathcal{I}; \Gamma \vdash P_{\mathtt{S}} \rhd \varnothing}\ \lfloor \mathrm{MACC} \rfloor$$

[8] By history-sensitivity A_i does not include any free state variable.

[9] ⌊BCH⌋/⌊SEL⌋ can be extended to delegation adding the following clause for $U_i = \langle \mathcal{L} \rangle$: (⌊BCH⌋) $\mathcal{C} \wedge A_i; \Gamma \vdash P_i \rhd \Delta, k[\mathtt{q}] : \mathcal{L}_i, x_i : \mathcal{L}$, and (⌊SEL⌋) $\mathcal{C} \wedge \underline{e_i}; \Gamma \vdash P[e_i'/x_i] \rhd \Delta', k[\mathtt{p}] : \mathcal{L}_j[e_i'/x_j]$ with $\Delta = \Delta'', e_i : \mathcal{L}_i'$ and $\Delta' = \Delta''$.

For readability we will simplify $\mathcal{I} \wedge \{\exists \mathtt{C}.\mathtt{credit}.\mathtt{C}.\mathtt{credit} \geqslant x\}$ with the equivalent predicate $\mathcal{I}$. Next, by $\lfloor$SEL$\rfloor$, setting $e = \mathtt{count} < 1000 \wedge x \geqslant 10$, and $E_{ok} = \mathtt{count} := \mathtt{count} + 1,\ \mathtt{credit} := \mathtt{credit} + x$:

$$\frac{\mathcal{I} \wedge e \supset (\mathtt{count} < 1000 \wedge y = \mathtt{count} \wedge E_{ok} = E_{ok} \wedge \mathcal{I}\,\mathtt{after}\,E_{ok})[\mathtt{count}/y] \quad \mathcal{I}; \Gamma \vdash \mathbf{0} \rhd z[\mathtt{S}] : \mathsf{end} \qquad \mathcal{I} \wedge \neg e \supset \mathsf{true} \quad \mathcal{I}; \Gamma \vdash \mathbf{0} \rhd z[\mathtt{S}] : \mathsf{end}}{\begin{array}{l}\mathcal{I}; \Gamma \vdash z[\mathtt{S}, \mathtt{C}]!\{e \mapsto \mathsf{ok}\langle \mathtt{count}\rangle(y)\langle E_{ok}\rangle.\mathbf{0}, \neg e \mapsto \mathsf{ko}.\mathbf{0}\} \\ \quad \vdash z[\mathtt{S}] : \mathtt{C}!\{\mathsf{ok}(y : \mathtt{Nat})\{\mathtt{count} < 1000 \wedge y = \mathtt{count}\}\langle E_{ok}\rangle.\mathsf{end},\ \mathsf{ko}.\mathsf{end}\}\end{array}}$$

where each line in the premise refers to a branch (i.e., ok and ko). The most delicate clause is $\mathcal{I} \wedge e \supset (\mathtt{count} < 1000 \wedge y = \mathtt{count} \wedge E_{ok} = E_{ok} \wedge \mathcal{I}\,\mathtt{after}\,E_{ok})[\mathtt{count}/y]$ which requires: (1) the interaction predicate to be satisfied under the current assumptions, and in fact $(\mathtt{count} < 1000 \wedge y = \mathtt{count})[\mathtt{count}/y]$ is implied by e, (2) the updates to be consistent, and in fact trivially $E_{ok} = E_{ok}$, and (3) the update to not invalidate the invariant, and in fact $\mathtt{credit} + x \geqslant 0 \wedge 0 \leqslant \mathtt{count} + 1 \leqslant 1000$ is true under the assumptions $\mathtt{credit} \geqslant 0$, $x \geqslant 10$ and $0 \leqslant \mathtt{count}$. Finally we apply $\lfloor$END$\rfloor$ to the second premise of each branch.

The effectiveness of the proof rules depends on the logic chosen for the predicates, which depends on the application scenario. An example which fits these criteria is the Presburger arithmetic, which is often sufficiently expressive: practical uses of multiplication are encodable [19], and formulae with quantifiers may be calculated efficiently [23, 25].

Proposition 10. *The proof of $\mathcal{C}; \Gamma \vdash N \rhd \Sigma$ is decidable if predicate evaluation is decidable.*

5 Soundness and completeness of the validation rules

We define a labelled transition relation for specifications $\langle \Gamma, \Sigma \rangle$ using the same labels as for networks. The main difference with the rules for networks is that predicates must be satisfied for the transition to occur. We illustrate below the most remarkable rules (the other rules are in Figure 9 in Appendix). The rule for session request:

$$\langle (a : \mathcal{G}, \Gamma); [\Delta]\sigma \rangle \xrightarrow{\overline{a}[\mathtt{n}]\langle s \rangle} \langle (a : \mathcal{G}, \Gamma); [\Delta, s[1] : \mathcal{G} \upharpoonright 1]\sigma \rangle$$

extends Δ with the new session, given that $a : \mathcal{G}$ in Γ and the current state satisfies assertion invariant A. The rule for session accept is dual. The rule for selection/send:

$$\frac{j \in I \quad \sigma \models A_j[\mathtt{n}/x_j] \quad \sigma' = \sigma\,\mathtt{after}\,E_j[\mathtt{n}/x_j]}{\langle \Gamma; [\Delta, s[\mathtt{p}] : \mathtt{q}!\{l_i(x_i : U_i)\{A_i\}\langle E_i\rangle.\mathcal{L}_i\}_{i \in I}]\sigma \rangle \xrightarrow{s[\mathtt{p},\mathtt{q}]!l_j\langle \mathtt{n} \rangle} \langle \Gamma; [\Delta, s[\mathtt{p}] : \mathcal{L}_j[\mathtt{n}/x_j]]\sigma' \rangle}$$

moves to the continuation $\mathcal{L}_j$ of the selected branch with the updated state σ', given that the sent value $\mathtt{n}$ satisfies predicate A_j for branch j in the current state σ.

Semantic conformance is defined using *conditional simulation* [6] to relate networks N to specifications $\langle \Gamma; \Sigma \rangle$.

Definition 11 (Conditional Simulation). A binary relation $\mathcal{R}$ over N and $\langle \Gamma; \Sigma \rangle$ is a *conditional simulation* if, for each $(N, \langle \Gamma; \Sigma \rangle) \in \mathcal{R}$, if $N \xrightarrow{\ell} N'$ with ℓ being:
(1) a branching then $\langle \Gamma; \Sigma \rangle$ is capable to move at the subject of ℓ, and if $\langle \Gamma; \Sigma \rangle \xrightarrow{\ell}$

$\langle \Gamma; \Sigma' \rangle$ then $(N', \langle \Gamma; \Sigma' \rangle) \in \mathcal{R}$;
(2) a select, session request/accept, τ then $\langle \Gamma; \Sigma \rangle \xrightarrow{\ell} \langle \Gamma; \Sigma' \rangle$ and $(N', \langle \Gamma; \Sigma' \rangle) \in \mathcal{R}$.
We write $N \lesssim \langle \Gamma; \Sigma \rangle$ if there exists a conditional simulation $\mathcal{R}$ s.t. $(N, \langle \Gamma; \Sigma \rangle) \in \mathcal{R}$.

Conditional simulation is like standard simulation for all types of actions except for branching, for which it requires N to be simulated only for legal values/labels (i.e., a process must conform to a given specification as long as its environment does so).

Definition 12 (Satisfaction). N satisfies Σ in Γ and $\mathcal{C}$, written $\mathcal{C}; \Gamma \models N \rhd \Sigma$, if for all closing substitutions $\tilde{\sigma}$ over N and Σ respecting Γ and $\mathcal{C}$, $N\tilde{\sigma} \lesssim \langle \Gamma; \Sigma\tilde{\sigma} \rangle$.

We write $\Gamma \models N \rhd \Sigma$ when $\mathcal{C}$ is true (e.g., for initial networks). Soundness and completeness for initial networks are stated below.

Theorem 13 (Soundness of Proof Rules). *Let N be an initial network. Then $\Gamma \vdash N \rhd \Sigma$ implies $\Gamma \models N \rhd \Sigma$.*

Theorem 14 (Completeness of Proof Rules). *Let $N \equiv \prod_{i \in I} [P_i]\sigma_{pi}$ be an initial network and $\Sigma = \prod_{i \in I} [\Delta_i]\sigma_{ai}$ be a specification. Assume that for all $i \in I$:* (1) $\mathtt{id}(\sigma_{pi}) = \mathtt{id}(\sigma_{ai})$, (2) $dom(\sigma_{pi}) = dom(\sigma_{ai})$, *and* (3) $\mathcal{I}(\sigma_{pi})$ *equivalent to* true. *If $\Gamma \models N \rhd \Sigma$ then $\Gamma \vdash N \rhd \Sigma$.*

(1-2) are for symmetry between N and Σ. (3) is necessary since the principals in N can make updates that differ from those made by the corresponding specifications in Σ; this may not compromise the observable behaviour of N with respect to Σ, but N may invalidate some principal invariant which would make the thesis false.

6 Related work and further topics

The preceding integrations of session types with logical constraints include [14], based on concurrent constraints ensuring bi-linear usage of channels, and [6], based on logical annotations on interactions, do not treat stateful properties. The combination of types and logical assertions referring to local state newly proposed in this paper enable fine-grained specifications and validation, which are not possible in [6, 14].

The expressiveness of the session type-based analyses has been greatly extended these past few years. On one side, the conversation calculus [9], contracts [12] and dynamic multirole session types [16] have opened the way to the modelling of protocols complex in their shapes, by describing more accurately how sessions can be joined or left, who is allowed participate. On the other side, works such as [6, 10] improved the way interactions inside a session are described: in [6], an assertion framework ensures logical properties on the communicated values, in [10], a security analysis guarantees that the coherence of the information flow is preserved. Our work improves the session type analyses in both directions: by proposing a division of the process being tested into separate principals that can join one or several sessions independently when conditions are matched and manage their own state, and by giving a description, inside each session, of the internal state of each participant and the property it should satisfy. A recent work [13] examines conditions to ensure that a stateful specification is robust w.r.t.

asynchronous communications. Our work provides a complete proof system ensuring soundness for processes, whereas [13] only addresses properties of types.

The refinement types for channels (e.g. [4]) specify value dependency with logical constraints. For example, one might write $?(x:\mathsf{int}, !\{y : \mathsf{int} \mid y > x\})$ (using the notation from [18]). It specifies a dependency at a *single point* (channel). Our theory, based on multiparty sessions, can verify processes against a contract globally agreed by multiple distributed peers. [3] uses refinement types for channels to verify authentication in multiparty session protocols, but does not consider multi-session properties.

The work [8] investigates a relationship between a dual intuitionistic linear logic and binary session types, and shows that the former defines a proof system for a session calculus which can automatically characterise and guarantee a session fidelity and global progress. None of the above works treat either virtual states or logical specifications for interleaved multiparty sessions.

The use of Rely-Guarantee conditions or other related methods [22] instead of a single invariant does not increase the expressiveness of our system, but could ease proofs for parallel composition.

A future direction is to link between our static analysis and a dynamic monitor-based approach. Using our local specification as a monitor at each end-point, incoming and outgoing messages can be verified and filtered. We are currently working on this topic with [24] based on the logic developed in this paper.

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A OOI case study: Instrument Command

To demonstrate our framework, we use a scenario based on the Instrument Command (IC) Usecase from the Ocean Observatories Initiative (OOI) [24]. The OOI is an NSF program to provide long-term infrastructure for delivering scientific data from a large network of ocean sensor systems to on-shore research stations around the US. Through the paper we will show how to validate process, engaging in multiple simultaneous instances of the type of session illustrated below. In the IC usecase, a user U obtains

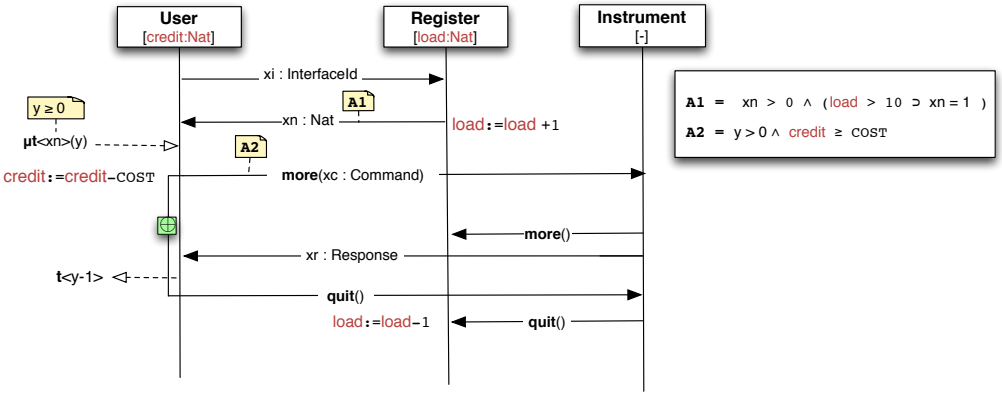

Fig. 5. Instrument control: an example of stateful specification

capabilities to use a particular instrument I from the service registry R. U initiates the session sending R an InterfaceId message which states the desired measurement type. R replies with the maximum number of measurements that U is permitted to make in this session. The process enters the main session loop; in each recursion step U has the choice of sending I the next Command via the more branch case, or to end the session via the quit case. The global specification declares that roles U and R must have credit and load state variables, respectively. credit is used to meter the usage of instruments by each principal, and load records the total current load of the instruments, which serves multiple users concurrently. Predicate A_1 ensures that R permits U to make $x_n > 0$ commands; however, if the current load is greater than a certain threshold (fixed as 10 in this example), then U is only permitted to make a single command. The subsequent update increments the load counter. Next, the recursion μt is annotated with a "loop counter" y, initialised to x_n, and the invariant predicate $y \geqslant 0$. The idea is that one command can be issued in each recursion step and C should not issue more than the permitted number of commands; the nested recursion variable t is accordingly annotated with $y - 1$. Within the recursion, the more message from U to I is guarded by the predicate that $y > 0$ and credit $\geqslant$ COST, where COST is the constant number of credits needed to perform one command; the associated update is to decrease credit by COST. The quit message to R at the end of the session has the effect of decrementing load.

A.1 Global assertion

$\mathcal{G}_{\text{IC}}$ is the global assertion for the protocol illustrated in Figure 5.

$$\begin{aligned}\mathcal{G}_{\text{IC}} = {} & \texttt{C} \to \texttt{R} : (x_i : \texttt{InterfaceId}).\\ & \texttt{R} \to \texttt{C} : (x_n : \texttt{int})\{x_n > 0 \wedge (\texttt{R.load} > 10 \supset x_n = 1)\}\langle \texttt{R.load} := \texttt{R.load} + 1\rangle.\\ & \mu \mathsf{t}\langle x_n\rangle(y : \texttt{int})\{y \geqslant 0\}.\texttt{C} \to \texttt{I} : \{\\ & \ \texttt{more}(x_c : \texttt{Command})\{y > 0 \wedge \texttt{C.credit} \geqslant \texttt{COST}\}\langle \texttt{C.credit} := \texttt{C.credit} - \texttt{COST}\rangle.\mathcal{G}_{\text{com}},\\ & \ \texttt{quit}().\mathcal{G}_{\text{end}}\}\end{aligned}$$

$$\mathcal{G}_{\text{com}} = \texttt{I} \to \texttt{C} : (x_r : \texttt{Response})\{\mathsf{true}\}.\texttt{I} \to \texttt{R} : \texttt{more}().\mathsf{t}\langle y - 1\rangle$$
$$\mathcal{G}_{\text{end}} = \texttt{I} \to \texttt{R} : \texttt{quit}()\langle \texttt{R.load} := \texttt{R.load} - 1\rangle.\mathsf{end}$$

$\mathcal{G}_{\text{IC}}$ specifies non-trivial dependencies between message behaviour and virtual state, which reflect the past and concurrent behaviours of the principal in other sessions. In the following discussion we focus on role R, hence we ignore the branch $\texttt{more}(x_c : \texttt{Command})$ which only concerns roles C and I.

A.2 Local assertion for R

$\mathcal{L}_{\texttt{R}}$ is the projection of $\mathcal{G}_{\text{IC}}$ in A.1 on role R. The projection makes use of a standard branch mergeability, which is an extension following e.g. [26].

$$\begin{aligned}\mathcal{L}_{\texttt{R}} &= \texttt{C}\,?(x_i : \texttt{InterfaceId}).\mathcal{L}'_{\texttt{R}}\\ \mathcal{L}'_{\texttt{R}} &= \texttt{C}\,!\,(x_n{:}\texttt{Nat})\{x_n > 0 \wedge (\texttt{load} > 10 \supset x_n = 1)\}\langle \texttt{load} := \texttt{load} + 1\rangle.\mathcal{L}''_{\texttt{R}}\\ \mathcal{L}''_{\texttt{R}} &= \mu \mathsf{t}\langle x_n\rangle(y : \texttt{Nat})\{y \geqslant 0\}.\texttt{I}\,?\{\texttt{more}().\mathsf{t}\langle y - 1\rangle, \texttt{quit}()\langle \texttt{load} := \texttt{load} - 1\rangle.\mathsf{end}\}\end{aligned}$$

Notice that the recursion invariant in $\mathcal{L}''_{\texttt{R}}$ would actually be, by definition, $\{\exists.\texttt{C.credit}.y \geqslant 0 \wedge \texttt{C.credit} \geqslant \texttt{COST}\}$.

A.3 Process

Process $\mathsf{P}_{\texttt{R}}$ accepts a request to engage in a session specified by global specification $\mathcal{G}_{\text{IC}}$ with the role of R. $\mathsf{P}_{\texttt{R}}$ is implementing the registry. The other roles involved in the session are user C, and instrument I. We omit the updates when empty, and the labels when there is only one branch.

$$\begin{aligned}\mathsf{P}_{\texttt{R}} &= a(z[\texttt{R}] : \mathcal{G}_{\text{IC}}).\mathsf{P}'_{\texttt{R}} \qquad \mathsf{P}'_{\texttt{R}} = z[\texttt{C},\texttt{R}]?(x_{id});\mathsf{P}''_{\texttt{R}}\\ \mathsf{P}''_{\texttt{R}} &= z[\texttt{R},\texttt{C}]!\{\{\texttt{load} > 10\} \mapsto \langle 1\rangle(x_1)\langle \texttt{load} := \texttt{load} + 1\rangle.\mathsf{P}''_{\texttt{R},1}\\ &\qquad\qquad \{\texttt{load} \leqslant 10\} \mapsto \langle 2\rangle(x_2)\langle \texttt{load} := \texttt{load} + 1\rangle.\mathsf{P}''_{\texttt{R},2}\\ \mathsf{P}''_{\texttt{R},n} &= (\mu X(y).z[\texttt{I},\texttt{R}]?\{\texttt{more}().t\langle y - 1\rangle,\ \texttt{quit}()\langle \texttt{load} := \texttt{load} - 1\rangle.\mathbf{0}\})\langle x_n\rangle\end{aligned}$$

In $\mathsf{P}'_{\texttt{R}}$, R receives the identifier x_{id} from C, and tests if state variable load is greater than the threshold of 10. If so, it sends the value 1; otherwise it sends 2. For brevity, we have parameterised the definitions of $\mathsf{P}''_{\texttt{R}}$ and $\mathsf{P}'''_{\texttt{R}}$ by $n \in \{1, 2\}$, with n used to initialise the later recursion parameter y. Each guarded-case leads to the appropriate $\mathsf{P}''_{\texttt{R},n}$, which increments load. R then enters the recursion, following I through the command-loop. R uses a branching inside the recursion: if more is received enters another recursion; otherwise, load is decremented.

A.4 Validation of two threads

We show selected extracts from the validation of a principal executing two parallel threads, each behaving as $P_{\mathtt{R}}$ in A.3. We set $\Gamma = a : \mathcal{G}_{\text{IC}}$, namely the principal can receive invitations to act in $\mathcal{G}_{\text{IC}}$. We set $\mathcal{I}(\sigma_p)$ ($\mathcal{I}$ for short) as $\mathtt{load} \geqslant 0$ and $\mathcal{C}$ initially true. We illustrate the validation of $[P_{\mathtt{R}} \mid P_{\mathtt{R}}]\sigma_p$ proceeding top-down and using the proof rules in Figure 4. The first rule to be applied is [N1]:

$$\frac{\sigma_p, \sigma_a \models \mathtt{load} \geqslant 0 \quad \mathtt{load} \geqslant 0 \wedge \mathsf{true}; \Gamma \vdash P_{\mathtt{R}} \mid P_{\mathtt{R}} \rhd \varnothing}{\mathsf{true}; \Gamma \vdash [P_{\mathtt{R}} \mid P_{\mathtt{R}}]\sigma_p \rhd [\varnothing]\sigma_a}$$

The process can be validated under the assumption that it runs in a state which satisfies the invariant $\mathtt{load} \geqslant 0$ (e.g., $\sigma_p(\mathtt{load}) = \sigma_a(\mathtt{load}) = 3$). Next we apply [PAR] which decomposes the derivation into the derivation of two identical threads: $\mathtt{load} \geqslant 0; \Gamma \vdash P_{\mathtt{R}} \rhd \varnothing$. We show only the derivation of one thread. The rule would apply in the same way even if the two parallel processes engaged in different types of conversations. Next we apply rule [MACC]:

$$\frac{\mathcal{I}; \Gamma \vdash \mathsf{P}'_{\mathtt{R}} \rhd z[\mathtt{R}] : \mathcal{L}_{\mathtt{R}}}{\mathcal{I}; \Gamma \vdash a[\mathtt{R}](z).\mathsf{P}'_{\mathtt{R}} \rhd \varnothing}$$

The premise is equivalent to $\mathcal{I}; \Gamma \vdash z[\mathtt{C}, \mathtt{R}]?(x_{id} : \mathtt{InterfaceId}).\mathsf{P}''_{\mathtt{R}} \rhd z[\mathtt{R}] : \mathcal{L}_{\mathtt{R}}$. After the application of rule $\lfloor$BCH$\rfloor$ we derive:

$$\frac{\begin{array}{c}\mathcal{I} \wedge \mathtt{load} > 10 \supset (A \wedge E = E \wedge \mathcal{I}\,\mathtt{after}\,E)[1/x_n] \quad \mathcal{I}; \mathsf{true}; \Gamma \vdash \mathsf{P}''_{\mathtt{R},1} \rhd z[\mathtt{R}] : \mathcal{L}''_{\mathtt{R}} \\ \mathcal{I} \wedge \mathtt{load} \leqslant 10 \supset A \wedge E = E \wedge \mathcal{I}\,\mathtt{after}\,E)[2/x_n] \quad \mathcal{I}; \mathsf{true}; \Gamma \vdash \mathsf{P}''_{\mathtt{R},2} \rhd z[\mathtt{R}] : \mathcal{L}''_{\mathtt{R}}\end{array}}{\mathcal{I}; \Gamma \vdash z[\mathtt{A}, \mathtt{C}]!\{\{\mathtt{load} > 10\} \mapsto \langle 1\rangle(x_n)\langle E\rangle.\mathsf{P}''_{\mathtt{R},1}, \{\mathtt{load} \leqslant 10\} \mapsto \langle 2\rangle(x_n)\langle E\rangle.\mathsf{P}''_{\mathtt{R},2}\} \rhd z[\mathtt{R}] : \mathcal{L}'_{\mathtt{R}}}$$

where we recall A is defined as $\{x_n > 0 \wedge (\mathtt{R.load} > 10 \supset x_n = 1)\}$ and E is $\langle \mathtt{load} := \mathtt{load} + 1\rangle$. Each branch in the premise can rely on the condition $\mathtt{load} > 10$ or $\neg(\mathtt{load} > 10)$. In the case $\mathtt{load} > 10$ it is necessary that x_n takes value 1; this is checked in the first premise for the first branch.

B Auxiliary Definitions

Definition 1 (Refinement). *A binary relation $\mathcal{R}$ over (Γ, Δ, σ) is a* refinement relation *if $(\Gamma_1, \Delta_1, \sigma_1)\mathcal{R}(\Gamma_2, \Delta_2, \sigma_2)$ implies one of the following conditions holds*

- $\langle \Gamma_1; [\Delta_1]\sigma_1\rangle \xrightarrow{\ell} \langle \Gamma_1; [\Delta'_1]\sigma'_1\rangle$ *with ℓ being a selection action then* $\langle \Gamma_2; [\Delta_2]\sigma_2\rangle \xrightarrow{\ell} \langle \Gamma_2; [\Delta_2]\sigma'_2\rangle$ *with* $(\Gamma_1, \Delta'_1, \sigma'_1)\mathcal{R}(\Gamma_2, \Delta'_2, \sigma'_2)$.
- $\langle \Gamma_2; [\Delta_2]\sigma_2\rangle \xrightarrow{\ell} \langle \Gamma_2; [\Delta'_2]\sigma'_2\rangle$ *with ℓ being a branching action, then* $\langle \Gamma_1; [\Delta_1]\sigma_1\rangle \xrightarrow{\ell} \langle \Gamma_1; [\Delta'_1]\sigma'_1\rangle$ *with* $(\Gamma_1, \Delta'_1, \sigma'_1)\mathcal{R}(\Gamma_2, \Delta'_2, \sigma'_2)$.

If $(\Gamma_1, \Delta_1, \sigma_1)\mathcal{R}(\Gamma_2, \Delta_2, \sigma_2)$ for some refinement relation $\mathcal{R}$, we say $(\Gamma_1, \Delta_1, \sigma_1)$ is a refinement *of $(\Gamma_2, \Delta_2, \sigma_2)$ (written $(\Gamma_1, \Delta_1, \sigma_1) \ni (\Gamma_2, \Delta_2, \sigma_2)$).*

Definition 2 (Projection). Assume $\mathtt{p}, \mathtt{q}, \mathtt{r} \in \mathcal{G}$ and $\mathtt{p} \neq \mathtt{q}$. The projection of $\mathcal{G}$ on $\mathtt{r} \in \mathcal{G}$, written $\mathcal{G} \upharpoonright \mathtt{r}$, is defined as follows.

$$(1)\ (\mathtt{p} \to \mathtt{q} : \{l_i(x_i : U_i)\{A_i\}\langle E_i\rangle.\mathcal{G}_i\}_{i\in I}) \upharpoonright \mathtt{r} = \begin{cases} \mathtt{q}!\{l_i(x_i : U_i)\{A_i\}\langle E_i \upharpoonright \mathtt{r}\rangle.(\mathcal{G}_i \upharpoonright \mathtt{p})\}_{i\in I} & \text{if } \mathtt{r} = \mathtt{p} \neq \mathtt{q}, \\ \mathtt{p}?\{l_i(x_i : U_i)\{A_i \upharpoonright \mathtt{r}\}\langle E_i \upharpoonright \mathtt{r}\rangle.(\mathcal{G}_i \upharpoonright \mathtt{q})\}_{i\in I} & \text{if } \mathtt{r} = \mathtt{q} \neq \mathtt{p}, \\ \mathcal{G}_1 \upharpoonright \mathtt{r} & \text{if } \mathtt{r} \neq \mathtt{q}, \mathtt{p} \end{cases}$$

$$(2)\ (\mathcal{G}_1 \mid \mathcal{G}_2) \upharpoonright \mathtt{r} = \begin{cases} \mathcal{G}_i \upharpoonright \mathtt{r} & \text{if } \mathtt{r} \in \mathcal{G}_i \text{ and } \mathtt{r} \notin \mathcal{G}_j, i \neq j \in \{1,2\} \\ \mathsf{end} & \text{if } \mathtt{r} \notin \mathcal{G}_1 \text{ and } \mathtt{r} \notin \mathcal{G}_2. \end{cases}$$

$$(3)\ (\mu \mathtt{t}\langle y : A'\rangle(x : S)\{A\}.\mathcal{G}) \upharpoonright \mathtt{r} = \begin{cases} \mu \mathtt{t}\langle y : A' \upharpoonright \mathtt{r}\rangle(x : S)\{A \upharpoonright \mathtt{r}\}.\mathcal{G} \upharpoonright \mathtt{r} & \text{if } \mathtt{r} \in \mathcal{G} \\ \mathsf{end} & \text{if } \mathtt{r} \notin \mathcal{G} \end{cases}$$

$$(4)\ \mathtt{t}\langle y : A'\rangle \upharpoonright \mathtt{r} = \mathtt{t}\langle y : A' \upharpoonright \mathtt{r}\rangle$$

B.1 Congruence, Reduction and Labelled Transitions

Figure 6 presents the full congruence rules for networks and processes, where the asynchronous messages are considered upon permutation. Figure 7 and Figure 9 illustrate the full transition rules for networks and specifications. Figure 7 models silent actions as reductions from Figure 8. In the paper we have represented silent actions explicitly in the LTS for a more concise presentation.

$$N \mid \varnothing \equiv N \qquad N_1 \mid N_2 \equiv N_1 \mid N_2 \qquad (N_1 \mid N_2) \mid N_3 \equiv N_1 \mid (N_2 \mid N_3) \text{ if } a \notin \mathsf{fn}(N)$$

$$(\nu s)\varnothing \equiv \varnothing \qquad (\nu s)(\nu s')N \equiv (\nu s')(\nu s)N \qquad (\nu s)N \mid N' \equiv (\nu s)(N \mid N') \quad \text{if } s \notin \mathsf{fn}(N)$$

$$s : h_1 \cdot (\mathtt{p}_1, \mathtt{p}_2, l\langle v\rangle) \cdot (\mathtt{q}_1, \mathtt{q}_2, l'\langle v'\rangle) \cdot h_2 \equiv s : h_1 \cdot (\mathtt{q}_1, \mathtt{q}_2, l'\langle v'\rangle) \cdot (\mathtt{p}_1, \mathtt{p}_2, l\langle v\rangle) \cdot h_2 \quad *\text{if } \mathtt{p}_1 \neq \mathtt{q}_1 \text{ or } \mathtt{p}_2 \neq \mathtt{q}_2$$

$$P \mid \mathbf{0} \equiv P \qquad P \mid Q \equiv Q \mid P \qquad (P \mid Q) \mid R \equiv P \mid (Q \mid R)$$

$$(\mu X(x).P)\langle e\rangle \equiv P[\mu X(x).P/X][e/x] \text{ where } X\langle e'\rangle[\mu X(x).P/X] \stackrel{\text{def}}{=} (\mu X(x).P)\langle e'\rangle$$

Fig. 6. Structural congruence for networks (top) and processes (bottom)

C Well-formed environments, kinding and typing

C.1 Well-formed environments

$$\frac{-}{\varnothing \vdash \mathit{Env}}\ [\textsc{Enul}] \qquad \frac{\Gamma \vdash U \rhd \mathit{Type} \quad x \notin \mathit{dom}(\Gamma)}{\Gamma, x : U \vdash \mathit{Env}}\ [\textsc{Esort}]$$

$$\frac{\Gamma \vdash S \triangleright \mathit{Type} \quad \{\Gamma \vdash \mathcal{L}_i \triangleright \mathit{Type}\}_{i=1..n} \quad X \notin \mathit{dom}(\Gamma)}{\Gamma, X : (S)\mathcal{L}_1@\mathtt{p}_1, \ldots, \mathcal{L}_n@\mathtt{p}_n \vdash \mathit{Env}}\ [\textsc{Erec}]$$

$$\frac{\Gamma \vdash \mathcal{G} \triangleright \mathit{Type} \quad a \notin \mathit{dom}(\Gamma)}{\Gamma, a : \mathcal{G} \vdash \mathit{Env}}\ [\textsc{Eshared}]$$

$$[\overline{a}[\mathtt{n}](y).P]\sigma \xrightarrow{\overline{a}[\mathtt{n}]\langle s\rangle} [P[s/y]]\sigma \quad [a[\mathtt{i}](y).P]\sigma \xrightarrow{a[i]\langle s\rangle} [P[s/y]]\sigma \quad (s \notin \mathsf{fn}(P))$$

$$[s[\mathtt{p},\mathtt{q}]!\{e_i \mapsto l_i\langle e_i'\rangle(x_i)\langle E_i\rangle; P_i\}_{i\in I}]\sigma \xrightarrow{s[\mathtt{p},\mathtt{q}]!l_j\langle v\rangle} [P[v/x_j]]\sigma' \quad (j \in I \quad \sigma \models e_j' \downarrow v \wedge e_j \quad \sigma' = \sigma\,\mathtt{after}\,E_j[v/x_j])$$

$$[s[\mathtt{p},\mathtt{q}]?\{l_i(x_i)\langle E_i\rangle.P_i\}_{i\in I}]\sigma \xrightarrow{s[\mathtt{p},\mathtt{q}]?l_j\langle v\rangle} [P_j[v/x_j]]\sigma' \quad (j \in I \quad \sigma' = \sigma\,\mathtt{after}\,E_j[v/x_j])$$

$$\frac{N \longrightarrow N'}{N \xrightarrow{\tau} N'} \qquad \frac{[P]\sigma \xrightarrow{\ell} [P']\sigma \quad (s \notin \mathsf{fn}(P))}{[\mathcal{E}[P]]\sigma \xrightarrow{\ell} [\mathcal{E}[P']]\sigma} \qquad \frac{N \equiv N_0 \quad N_0 \xrightarrow{\ell} N_0' \quad N_0' \equiv N' \quad \mathsf{fn}(\ell) \notin bn(\mathcal{E}[\_])}{\mathcal{E}[N] \xrightarrow{\ell} \mathcal{E}[N']}$$

Fig. 7. Labelled transition for networks

$$[\overline{a}[\mathtt{n}](y).P_1 \mid Q_1]\sigma_1 \mid \textstyle\prod_{2\leqslant i\leqslant n}[a[\mathtt{i}](y_i).P_i \mid Q_i]\sigma_i \longrightarrow (\nu s)(s : \varnothing \mid \textstyle\prod_{1\leqslant i\leqslant n}[P_i[s/y_i] \mid Q_i]\sigma_i)$$

$$\frac{j \in I \qquad \sigma \models e_j \qquad \sigma \models e_j' \downarrow v \qquad \sigma' = \sigma\,\mathtt{after}\,E_j[v/x_j]}{[s[\mathtt{p},\mathtt{q}]!\{e_i \mapsto l_i\langle e_i'\rangle(x_i)\langle E_i\rangle; P_i\}_{i\in I} \mid Q]\sigma \mid s : h \longrightarrow [P_j[v/x_j] \mid Q]\sigma' \mid s : h \cdot (\mathtt{p},\mathtt{q}, l_j\langle v\rangle)}$$

$$\frac{j \in I \qquad \sigma' = \sigma\,\mathtt{after}\,E_j[v/x_j]}{[s[\mathtt{p},\mathtt{q}]?\{l_i(x_i)\langle E_i\rangle.P_i\}_{i\in I} \mid Q]\sigma \mid s : (\mathtt{p},\mathtt{q}, l_j\langle v\rangle) \cdot h \longrightarrow [P_j[v/x_j] \mid Q]\sigma' \mid s : h}$$

$$\frac{P \equiv P_0 \quad P_0 \longrightarrow P_0' \quad P_0' \equiv P'}{P \longrightarrow P'} \qquad \frac{N \equiv N_0 \quad N_0 \longrightarrow N_0' \quad N' \equiv N_0'}{\mathcal{E}[N] \longrightarrow \mathcal{E}[N']}$$

Fig. 8. Reduction for networks

C.2 Kinding system

Type

$$\frac{\Gamma \vdash \mathit{Env}}{\Gamma \vdash \mathit{Type}} \quad \lfloor \textsc{kBase} \rfloor$$

Value Types

$$\frac{\Gamma \vdash \mathcal{L} \triangleright \mathit{Type} \qquad \mathtt{ftv}(\mathcal{L}) = \varnothing}{\Gamma \vdash \langle\mathcal{L}\rangle \triangleright \mathit{Type}} \quad \lfloor \textsc{kMar} \rfloor$$

$$\frac{\Gamma \vdash \mathit{Env}}{\Gamma \vdash \mathtt{int} \triangleright \mathit{Type}} \quad \lfloor \textsc{kInt} \rfloor \qquad \frac{\Gamma \vdash \mathit{Env}}{\Gamma \vdash \mathtt{bool} \triangleright \mathit{Type}} \quad \lfloor \textsc{kBool} \rfloor$$

$$\frac{\Gamma \vdash \mathit{Env}}{\Gamma \vdash \mathtt{string} \triangleright \mathit{Type}} \quad \lfloor \textsc{kStr} \rfloor \qquad \frac{\Gamma \vdash \mathit{Env}}{\Gamma \vdash \mathtt{nat} \triangleright \mathit{Type}} \quad \lfloor \textsc{kNat} \rfloor$$

Session Types (global)

$$\frac{\{\Gamma \vdash i : \mathtt{nat} \quad \Gamma, x_i : U_i \vdash \mathcal{G}_i, A_i, E_i, U_i \triangleright \mathit{Type} \quad x_i \notin dom(\Gamma) \quad \Gamma \vdash l_i : \mathtt{string}\}_{i\in I} \quad l_i \neq l_j \ \text{ if } i \neq j \text{ for all } i,j \in I}{\Gamma \vdash \mathtt{p} \to \mathtt{q} : \{l_i(x_i : U_i)\{A_i\}\langle E_i\rangle.\mathcal{G}_i\}_{i\in I} \triangleright \mathit{Type}} \quad \lfloor \textsc{kGSnd} \rfloor$$

$$\frac{\Gamma \vdash \mathcal{G}_1 \triangleright \mathit{Type} \qquad \Gamma \vdash \mathcal{G}_2 \triangleright \mathit{Type}}{\Gamma \vdash \mathcal{G}_1 \mid \mathcal{G}_2 \triangleright \mathit{Type}} \quad \lfloor \textsc{kGPar} \rfloor \qquad \frac{\Gamma \vdash A', A, S \triangleright \mathit{Type} \quad \Gamma, x : S \vdash \mathcal{L} \triangleright \mathit{Type}}{\Gamma \vdash \mu \mathsf{t}\langle y : A'\rangle(x : S)\{A\}.\mathcal{L} \triangleright \mathit{Type}} \quad \lfloor \textsc{kRec} \rfloor$$

$$\frac{\Gamma \vdash A' \triangleright \mathit{Type} \qquad \Gamma \vdash \mathit{Env}}{\Gamma \vdash \mathsf{t}\langle y : A'\rangle \triangleright \mathit{Type}} \quad \lfloor \textsc{kGCall} \rfloor \qquad \frac{\Gamma \vdash \mathit{Env}}{\Gamma \vdash \mathsf{end} \triangleright \mathit{Type}} \quad \lfloor \textsc{kGEnd} \rfloor$$

$$\frac{\langle \Gamma; \Sigma_1\rangle \xrightarrow{\ell} \langle \Gamma; \Sigma_1'\rangle}{\langle \Gamma; \Sigma_1, \Sigma_2\rangle \xrightarrow{\ell} \langle \Gamma; \Sigma_1', \Sigma_2\rangle} \quad \frac{\langle \Gamma; \Sigma_2\rangle \xrightarrow{\ell} \langle \Gamma; \Sigma_2'\rangle}{\langle \Gamma; \Sigma_1, \Sigma_2\rangle \xrightarrow{\ell} \langle \Gamma; \Sigma_1, \Sigma_2'\rangle} \quad \lfloor \text{TR-CTX1/TR-CTX2} \rfloor$$

$$\langle \Gamma; \Sigma\rangle \xrightarrow{\tau} \langle \Gamma; \Sigma\rangle \quad \lfloor \text{TR-TAU} \rfloor$$

$$\langle (a : \mathcal{G}, \Gamma); [\Delta]\sigma\rangle \xrightarrow{\overline{a}[\mathtt{n}]\langle s\rangle} \langle (a : \mathcal{G}, \Gamma); [\Delta, s[\mathtt{1}] : \mathcal{G} \upharpoonright \mathtt{1}]\sigma\rangle \quad \lfloor \text{TR-A-MREQ} \rfloor$$

$$\langle (a : \mathcal{G}, \Gamma); [\Delta]\sigma\rangle \xrightarrow{a[i]\langle s\rangle} \langle (a : \mathcal{G}, \Gamma); [\Delta, s[\mathtt{i}] : \mathcal{G} \upharpoonright \mathtt{i}]\sigma\rangle \quad \lfloor \text{TR-A-MACC} \rfloor$$

$$\frac{j \in I \qquad \sigma \models A[\mathtt{n}/x_j] \qquad \sigma' = \sigma\,\mathtt{after}\,E_j[\mathtt{n}/x_j]}{\langle \Gamma; [\Delta, s[\mathtt{p}] : \mathtt{q}?\{l_i(x_i : U_i)\{A_i\}\langle E_i\rangle.\mathcal{L}_i\}_{i\in I}]\sigma\rangle \xrightarrow{s[\mathtt{q},\mathtt{p}]?l_j\langle \mathtt{n}\rangle} \langle \Gamma; [\Delta, s[\mathtt{p}] : \mathcal{L}_j[\mathtt{n}/x_j]]\sigma'\rangle} \quad \lfloor \text{TR-A-BCH} \rfloor$$

$$\frac{j \in I \qquad \sigma \models A[\mathtt{n}/x_j] \qquad \sigma' = \sigma\,\mathtt{after}\,E_j[\mathtt{n}/x_j]}{\langle \Gamma; [\Delta, s[\mathtt{p}] : \mathtt{q}!\{l_i(x_i : U_i)\{A_i\}\langle E_i\rangle.\mathcal{L}_i\}_{i\in I}]\sigma\rangle \xrightarrow{s[\mathtt{p},\mathtt{q}]!l_j\langle \mathtt{n}\rangle} \langle \Gamma; [\Delta, s[\mathtt{p}] : \mathcal{L}_j[\mathtt{n}/x_j]]\sigma'\rangle} \quad \lfloor \text{TR-A-SEL} \rfloor$$

$$\frac{j \in I \qquad \sigma \models A_j \qquad \sigma' = \sigma\,\mathtt{after}\,E_j \quad U_j = \langle \mathcal{L}\rangle}{\langle \Gamma; [\Delta, s[\mathtt{p}] : \mathtt{q}?\{l_i(x_i : U_i)\{A_i\}\langle E_i\rangle.\mathcal{L}_i\}_{i\in I}]\sigma\rangle \xrightarrow{s[\mathtt{q},\mathtt{p}]?l_j\langle t[\mathtt{r}]\rangle} \langle \Gamma; [\Delta, s[\mathtt{p}] : \mathcal{L}_j, t[\mathtt{r}] : \mathcal{L}]\sigma'\rangle} \quad \lfloor \text{TR-A-DELBCH} \rfloor$$

$$\frac{j \in I \qquad \sigma \models A_j \qquad \sigma' = \sigma\,\mathtt{after}\,E_j \quad U_j = \langle \mathcal{L}\rangle}{\langle \Gamma; [\Delta, s[\mathtt{p}] : \mathtt{q}!\{l_i(x_i : U_i)\{A_i\}\langle E_i\rangle.\mathcal{L}_i\}_{i\in I}, t[\mathtt{r}] : \mathcal{L}]\sigma\rangle \xrightarrow{s[\mathtt{p},\mathtt{q}]!l_j\langle t[\mathtt{r}]\rangle} \langle \Gamma; [\Delta, s[\mathtt{p}] : \mathcal{L}_j]\sigma'\rangle} \quad \lfloor \text{TR-A-DELSEL} \rfloor$$

Fig. 9. Labelled transition for specifications

Session Types (local)

$$\frac{\{\Gamma \vdash i : \mathtt{nat} \quad \Gamma, x_i : U_i \vdash \mathcal{L}_i, A_i, E_i, U_i \blacktriangleright \mathit{Type} \quad x_i \notin \mathit{dom}(\Gamma) \quad \Gamma \vdash l_i : \mathtt{string}\}_{i\in I} \quad l_i \neq l_j \text{ if } i \neq j \text{ for all } i, j \in I)}{\Gamma \vdash \mathtt{p}!\{l_i(x_i : U_i)\{A_i\}\langle E_i\rangle.\mathcal{L}_i\}_{i\in I} \blacktriangleright \mathit{Type}} \quad \lfloor \text{KLSND} \rfloor$$

$$\frac{\{\Gamma \vdash i : \mathtt{nat} \quad \Gamma, x_i : U_i \vdash \mathcal{L}_i, A_i, E_i, U_i \blacktriangleright \mathit{Type} \quad x_i \notin \mathit{dom}(\Gamma) \quad \Gamma \vdash l_i : \mathtt{string} \quad l_i \neq l_j \text{ if } i \neq j \text{ for all } i, j \in I)\}_{i\in I}}{\Gamma \vdash \mathtt{p}?\{l_i(x_i : U_i)\{A_i\}\langle E_i\rangle.\mathcal{L}_i\}_{i\in I} \blacktriangleright \mathit{Type}} \quad \lfloor \text{KLRCV} \rfloor$$

$$\frac{\Gamma \vdash A', A, S \blacktriangleright \mathit{Type} \quad \Gamma, x : S \vdash \mathcal{L} \blacktriangleright \mathit{Type}}{\Gamma \vdash \mu \mathsf{t}\langle y : A'\rangle(x : S)\{A\}.\mathcal{L} \blacktriangleright \mathit{Type}} \quad \lfloor \text{KLREC} \rfloor$$

$$\frac{\Gamma \vdash A' \blacktriangleright \mathit{Type} \qquad \Gamma \vdash \mathit{Env}}{\Gamma \vdash \mathsf{t}\langle y : A'\rangle \blacktriangleright \mathit{Type}} \quad \lfloor \text{KLCALL} \rfloor \qquad \frac{\Gamma \vdash \mathit{Env}}{\Gamma \vdash \mathsf{end} \blacktriangleright \mathit{Type}} \quad \lfloor \text{KLEND} \rfloor$$

Specifications

$$\frac{\Gamma \vdash \mathcal{G} \rhd \mathit{Type} \qquad \mathcal{L} = \mathcal{G} \upharpoonright \mathtt{p} \qquad \Gamma \vdash \mathit{Env} \quad s \notin \mathit{dom}(\Gamma) \qquad \Gamma \vdash \Delta \rhd \mathit{Type}}{\Gamma \vdash \Delta, s[\mathtt{p}] : \mathcal{L} \blacktriangleright \mathit{Type}} \quad \lfloor \text{DSES} \rfloor$$

$$\frac{\Gamma \vdash \mathit{Env}}{\Gamma \vdash \varnothing \blacktriangleright \mathit{Type}} \quad \lfloor \text{DNULL} \rfloor$$

Predicates, expressions, updates

$$\frac{\Gamma \vdash \mathit{Env}}{\Gamma \vdash \mathsf{true}, \mathsf{false} \triangleright \mathit{Type}} \ \lfloor \textsc{pBasic} \rfloor$$

$$\frac{\Gamma \vdash A \triangleright \mathit{Type}}{\Gamma \vdash \neg A \triangleright \mathit{Type}} \ \lfloor \textsc{pNeg} \rfloor \qquad \frac{i = 1,2 \quad \Gamma \vdash A_i \triangleright \mathit{Type}}{\Gamma \vdash A_1 \wedge A_2 \triangleright \mathit{Type}} \ \lfloor \textsc{pAnd} \rfloor \qquad \frac{\Gamma, x : S \vdash A \triangleright \mathit{Type}}{\Gamma \vdash \exists x.A \triangleright \mathit{Type}} \ \lfloor \textsc{pEx} \rfloor$$

$$\frac{\Gamma \vdash \mathit{Env} \qquad i = 1,2 \ \ \Gamma \vdash e_i : \mathtt{int/nat}}{\Gamma \vdash e_1 > e_2 \triangleright \mathit{Type}} \ \lfloor \textsc{Exp1} \rfloor$$

$$\frac{\Gamma \vdash \mathit{Env} \qquad i = 1,2 \ \ \Gamma \vdash e_i : \mathtt{int} \ (\text{same for } \mathtt{nat/string/bool})}{\Gamma \vdash e_1 = e_2 \triangleright \mathit{Type}} \ \lfloor \textsc{Exp2} \rfloor$$

$$\frac{\Gamma \vdash \mathit{Env} \qquad i = 1,2 \ \ \Gamma \vdash e_i : \mathtt{bool} \qquad \mathbf{op} \in \{\wedge, \vee\}}{\Gamma \vdash e_1 \mathbf{op}\ e_2 \triangleright \mathit{Type}} \ \lfloor \textsc{Exp3} \rfloor$$

$$\frac{\Gamma \vdash \mathit{Env} \qquad i = 1,2 \ \ \Gamma \vdash e_i : \mathtt{int} \qquad \mathbf{op} \in \{+, -\}}{\Gamma \vdash e_1 \mathbf{op}\ e_2 \triangleright \mathit{Type}} \ \lfloor \textsc{Exp4} \rfloor$$

$$\frac{\Gamma \vdash \mathit{Env}}{\Gamma \vdash \varnothing \triangleright \mathit{Type}} \ \lfloor \textsc{NoUpd} \rfloor \qquad \frac{\Gamma \vdash \mathit{Env} \qquad \Gamma \vdash E \triangleright \mathit{Type} \quad \Gamma \vdash x : S \qquad \Gamma \vdash e : S}{\Gamma \vdash x := e; E \triangleright \mathit{Type}} \ \lfloor \textsc{Upd} \rfloor$$

D Complexity of well-assertedness and proof rules

D.1 Checking History Sensitivity

History-sensitivity can be checked by inductive rules (see 10). The syntactic checker uses the environment $\mathcal{E}$ defined by the following grammar:

$$\mathcal{E} := \varnothing \mid \mathcal{E}, x@\mathtt{p} \mid \mathcal{E}, x@\mathtt{L} \mid \mathcal{E}, \mathsf{t} : x@\mathtt{L}$$

Expressions of the form $y@\mathtt{p}$ assign a state variable y to a role, and $y@\mathtt{L}$ assigns an interaction variable or recursion parameters to a *location* $\mathtt{L}$. A location is a set $\{\mathtt{p}, \mathtt{p}'\}$ of the two roles who know x. The checker relies on the annotation of the recursion parameters with their locations and we assume recursion parameter to be known by only two roles. We denote the domain of $\mathcal{E}$ with $dom(\mathcal{E})$. We write $\mathcal{E} \vdash x@\mathtt{p}$ when $\mathtt{p} = \mathcal{E}(x)$, $\mathtt{p} \in \mathcal{E}(x)$, or $\mathcal{E}(\mathsf{t}) = x@\mathtt{L}$ and $\mathtt{p} \in \mathtt{L}$.

$$\frac{\forall i \in I, \quad \mathcal{E}, x_i@\{\mathtt{p},\mathtt{q}\} \vdash \mathcal{G}_i \quad \forall y \in (var(A_i) \cup var(E_i))\backslash x_i,\ \mathcal{E} \vdash y@\mathtt{p} \quad \forall y \in var(E_i)\backslash x_i,\ \mathcal{E} \vdash y@\mathtt{q}}{\mathcal{E} \vdash \mathtt{p} \to \mathtt{q} : \{l_i(x_i)\{A_i\}\langle E_i\rangle.\mathcal{G}_i\}_{i \in I}}$$

$$\frac{\mathcal{E} \vdash \mathcal{G} \quad \mathcal{E} \vdash \mathcal{G}'}{\mathcal{E} \vdash \mathcal{G}, \mathcal{G}'} \qquad \frac{}{\mathcal{E} \vdash \mathsf{end}} \qquad \frac{\forall y \in var(e),\ \forall \mathtt{r} \in \mathtt{L},\ \mathcal{E} \vdash y@\mathtt{r}}{\mathcal{E}, \mathsf{t} : x@\mathtt{L} \vdash \mathsf{t}\langle e\rangle} \qquad \frac{\mathcal{E}, \mathsf{t} : x@\mathtt{L} \vdash \mathcal{G} \quad dom(\mathcal{E}) \supseteq var(A)\backslash x}{\mathcal{E} \vdash \mu \mathsf{t}\langle e\rangle(x@\mathtt{L})\{A\}.\mathcal{G}}$$

Fig. 10. Syntactic checker for history-sensitivity

The rules in Figure 10 enforce well-assertedness by restricting the set of variables that can be used in each predicate and update. The first rule requires that each role knows all the interaction variables of the predicate to be checked at its side. The other rules are straightforward.

D.2 Proof of Proposition 5

We recall the statement of Proposition 5. Given a global assertion $\mathcal{G}$, let m be the size of the syntactic tree of $\mathcal{G}$, n be the maximum number of variables occurring in each predicate in $\mathcal{G}$, and $eval(A)$ be the complexity of predicate evaluation. History-sensitivity for $\mathcal{G}$ can be checked in $O(m \times n)$. Temporal-satisfiability for $\mathcal{G}$ is decidable if predicate evaluation is decidable; if decidable temporal-satisfiability can be checked in $O(m) \times eval(A)$.

The size m of $\mathcal{G}$ is defined by the following function, by induction:

- $size(\mathtt{p} \to \mathtt{q} : \{l_i(x_i : U_i)\{A_i\}\langle E_i\rangle.\mathcal{G}_i\}_{i \in I}) = 1 + \sum_{i=1}^{n} size(\mathcal{G}_i)$
- $size(\mathcal{G}_1 \mid \mathcal{G}_2) = 1 + size(\mathcal{G}_1) + size(\mathcal{G}_2)$
- $size(\mu \mathsf{t}\langle y : A'\rangle(x : S)\{A\}.\mathcal{G}) = 1 + size(\mathcal{G})$
- $size(\mathsf{t}\langle y : A'\rangle) = 1$
- $size(\mathsf{end}) = 1$

Proof. For history-sensitivity, the size of a proof tree for $\mathcal{G}$ obtained using the rules in Figure 10 has the same order of magnitude of the size of the syntactic tree of $\mathcal{G}$, hence $O(m)$, and at each point of the proof tree only a purely syntactic check on the variables appearing in the current predicate will occur, which are at most n. For temporal-satisfiability, the number of invocation of $\mathtt{ts}$ is $O(m)$ and at each invocation at most one predicate must be evaluated with complexity $eval(A)$ (if decidable).

D.3 Proof of Proposition 10 (complexity of proof rules)

We recall the statement of Proposition 10. The proof of $\mathcal{C}; \Gamma \vdash N \rhd \Sigma$ is decidable if predicate evaluation is decidable.

Proof. The typing rules in [20] enable decidable validation. The proof rules in Figure 4 have the same structure (i.e., they decompose the validation in the same way) as the rules in [20], to which they add the evaluation of (a linear number of) predicates. If predicate evaluation is decidable, the proof tree has depth linear with respect to the syntactic tree of P.

E Message Assertions

We introduce the definitions for processes with queues. The aim is to take into account, in the proof of soundness of the validation rules, the mechanisms of message exchange of runtime processes. We use *message assertions* which abstract messages in queues.

Definition 15 (Message Assertions). The syntax of endpoint assertions is extended as follows:

$$\mathcal{L} ::= \ldots \mid \mathcal{M} \mid \mathcal{M}; \mathcal{L} \qquad\qquad \mathcal{M} ::= \mathtt{p}!l\langle v\rangle \mid \mathcal{M}; \mathcal{M}'$$

We call $\mathcal{M}$ a *message assertion*.

In Definition 15, $\mathtt{p}!l\langle v\rangle$ represents a label/value $l\langle v\rangle$ in the queue for participant $\mathtt{p}$, and $\mathcal{M}; \mathcal{M}'$ represents a queue with multiple elements.

Figure 11 presents the additional validation rules for runtime processes (to extend the rules in Figure 4).

Figure 12 presents the additional transition rules for message assertions.

$$\frac{\mathcal{C}; \Gamma \vdash N \rhd [\Delta_1, s[\mathtt{1}] : \mathcal{L}_1]\sigma_1, \ldots, [\Delta_n, s[\mathtt{n}] : \mathcal{L}_n]\sigma_n \qquad \{s[\mathtt{i}] : \mathcal{L}_i\}_{1 \geqslant i \geqslant n} \text{ coherent}}{\mathcal{C}; \Gamma \vdash (\nu s)N \rhd [\Delta_1]\sigma_1, \ldots, [\Delta_n]\sigma_n} \lfloor \textsc{Cres} \rfloor$$

$$\mathcal{C}; \Gamma \vdash s : \varnothing \rhd \{s[\mathtt{i}] : \varnothing\}_{1 \geqslant i \geqslant n} \lfloor \textsc{Qnil} \rfloor$$

$$\frac{\mathcal{C}; \Gamma \vdash s : h \rhd [\Delta, s[\mathtt{p}] : \mathcal{L}]\sigma, \Sigma}{\mathcal{C}; \Gamma \vdash s : h \cdot (\mathtt{p}, \mathtt{q}, l\langle v \rangle) \rhd [\Delta, s[\mathtt{p}] : \mathtt{q}!l\langle v \rangle; \mathcal{L}]\sigma, \Sigma} \lfloor \textsc{Qval} \rfloor$$

Fig. 11. Additional Proof Rules for Runtime Networks and Processes

$$\frac{-}{\langle \Gamma, [\Delta, s[\mathtt{p}] : \mathtt{q}!l\langle v \rangle; \mathcal{L}]\sigma \rangle \xrightarrow{s[\mathtt{p},\mathtt{q}]!l\langle v \rangle} \langle \Gamma, [\Delta, s[\mathtt{p}] : \mathcal{L}]\sigma \rangle} \lfloor \text{T1} \rfloor$$

$$\frac{j \in I \qquad \sigma \models A_j[\mathtt{n}/x_j] \downarrow \mathsf{true} \quad \sigma' = \sigma \,\mathtt{after}\, E_j[\mathtt{n}/x_j]}{\langle \Gamma; [s[\mathtt{p}] : \mathtt{q}!\{l_i(x_i : U_i)\{A_i\}\langle E_i \rangle.\mathcal{L}_i\}_{i \in I}]\sigma \rangle \xrightarrow{\tau} \langle \Gamma; [s[\mathtt{p}] : \mathtt{q}!l_j\langle \mathtt{n} \rangle; \mathcal{L}_j[v/x_j]]\sigma' \rangle} \lfloor \text{T2} \rfloor$$

$$\frac{j \in I \qquad \sigma \models A_j[\mathtt{n}/x_j] \downarrow \mathsf{true} \quad \sigma'' = \sigma' \,\mathtt{after}\, E_j[\mathtt{n}/x_j]}{\begin{array}{r}\langle \Gamma; [s[\mathtt{p}] : \mathtt{q}!l_j\langle \mathtt{n} \rangle; \mathcal{L}]\sigma, [s[\mathtt{q}] : \mathtt{p}?\{l_i(x_i : U_i)\{A_i\}\langle E_i \rangle.\mathcal{L}_i\}_{i \in I}]\sigma' \rangle \\ \xrightarrow{\tau} \langle \Gamma; [s[\mathtt{p}] : \mathcal{L}]\sigma, [\mathcal{L}_j[\mathtt{n}/x_j]]\sigma'' \rangle\end{array}} \lfloor \text{T3} \rfloor$$

$$\frac{j \in I \qquad \sigma \models A_j \downarrow \mathsf{true} \quad \sigma' = \sigma \,\mathtt{after}\, E_j \quad U_j = \langle \mathcal{L} \rangle}{\begin{array}{r}\langle \Gamma; [s[\mathtt{p}] : \mathtt{q}!\{l_i(x_i : U_i)\{A_i\}\langle E_i \rangle.\mathcal{L}_i\}_{i \in I}, t[\mathtt{r}] : \mathcal{L}]\sigma \rangle \\ \xrightarrow{\tau} \langle \Gamma; [s[\mathtt{p}] : \mathtt{q}!l_j\langle t[\mathtt{r}] \rangle; \mathcal{L}_j]\sigma' \rangle\end{array}} \lfloor \text{T4} \rfloor$$

$$\frac{j \in I \qquad \sigma \models A_j \downarrow \mathsf{true} \quad \sigma'' = \sigma' \,\mathtt{after}\, E_j \quad U_j = \langle \mathcal{L} \rangle}{\begin{array}{r}\langle \Gamma; [s[\mathtt{p}] : \mathtt{q}!l_j\langle v \rangle; \mathcal{L}', t[\mathtt{r}] : \mathcal{L}]\sigma, [s[\mathtt{q}] : \mathtt{p}?\{l_i(x_i : U_i)\{A_i\}\langle E_i \rangle.\mathcal{L}_i\}_{i \in I}]\sigma' \rangle \\ \xrightarrow{\tau} \langle \Gamma; [s[\mathtt{p}] : \mathcal{L}']\sigma, [\mathcal{L}_j, t[\mathtt{r}] : \mathcal{L}]\sigma'' \rangle\end{array}} \lfloor \text{T5} \rfloor$$

Fig. 12. Labelled transition for message assertions

F Soundness

F.1 Auxiliary Lemmas

This section contains auxiliary lemmas for soundness. The proofs of Lemma 1, Lemma 2 can be found below. The proofs of Lemma 3 and Lemma 4 are similar to the ones in [6] (hence omitted) as the stated properties do not directly involve the state.

Substitution The substitution lemma uses the following lemma saying that any substitution of a free variable with a value in a local assertion preserves well-assertedness.

Lemma 16. *Let $\mathcal{L}$ be a well-asserted local assertion (and well-typed wrt the underlying typing discipline), $x : U$ be an interaction variable, $v : U$ be a value of the same type as x. If $\mathcal{C}[v/x]$ admits solutions then $\mathcal{L}[v/x]$ is well-asserted.*

Proof. History sensitivity is clearly not affected by the substitution of an interaction variable with a value, as it is based on the notion of knowledge and a value is obviously known by any participant. For invariant stability, assume $\mathcal{L}[v/x]$. Since $\mathcal{L}$ is temporal satisfiable, by hypothesis the checker will return false for $\mathcal{L}[v/x]$ because of the otherwise case is met in (1) or (2). In both cases, if the predicate $(A_{inv} \wedge A_{bag} \wedge A_i \supset A_{inv}\ \texttt{after}\ E_i)[v/x]$ is false then also the its (stronger) unsubstituted version is false, which makes $\mathcal{L}$ not invariant stable contradicting the hypothesis.

Lemma 1 (Substitution). *Let $\mathcal{I}; \mathcal{C}; \Gamma \vdash P \rhd \Delta$ with Δ well-asserted and $x : U$ be an interaction variable and $v : U$ be a value. If $x \in \mathsf{fn}(P)$ then $\mathcal{C}[v/x]; \Gamma \vdash P[v/x] \rhd \Delta[v/x]$ and $\Delta[v/x]$ is well-asserted.*

Proof. The proof is by on the validation rules. We proceed by case analysis o the rules in Figure 4.

Case $\lfloor$SEL$\rfloor$. We set $P = s[\mathtt{p}, \mathtt{q}]!\{e_i \mapsto l_i\langle e_i'\rangle(x_i)\langle E_i\rangle; P_i\}_{i\in I}$ and $\Delta = \Delta', s[\mathtt{p}] : \mathtt{q}!\{l_j(x_j : U_j)\{A_j\}\langle E_j\rangle.\mathcal{L}_j\}_{j\in J}$. We first assume $x : S$. By $\lfloor$SEL$\rfloor$:

$$\frac{\begin{array}{l}\forall i \in I \exists j \in J\ l_i = l_j \quad \mathcal{C} \wedge e_i \supset (A_j \wedge (E_i = E_j) \wedge \mathcal{C}\ \texttt{after}\ E_i)[e_i'/x_i] \\ \mathcal{C} \wedge \underline{e_i}; \Gamma \vdash P_i[e_i'/x_i] \rhd \Delta', s[\mathtt{p}] : \mathcal{L}_j[e_i'/x_j]\end{array}}{\mathcal{C}; \Gamma \vdash P \rhd \Delta', s[\mathtt{p}] : \mathtt{q}!\{l_j(x_j : U_j)\{A_j\}\langle E_j\rangle.\mathcal{L}_j\}_{j\in J}} \quad (1)$$

Without loss of generality we assume $x \notin \{x_j\}_{j\in J}$. The first premise of (1) entails the following predicate

$$(\mathcal{C} \wedge e_i \supset (A_j \wedge (E_i = E_j) \wedge \mathcal{C}\ \texttt{after}\ E_i)[e_i'/x_j])[v/x] \quad (2)$$

since the free occurrences of x (if any) in the first premise of (1) are supposed to be universally quantified. By definition, (2) is equivalent to

$$(\mathcal{C} \wedge e_i)[v/x] \supset (A_j \wedge (E_i = E_j) \wedge \mathcal{C}\ \texttt{after}\ E_i)[e_i'/x_j][v/x] \quad (3)$$

Moreover, by inductive hypothesis, we have

$$\mathcal{C}[v/x]; \Gamma \vdash P_i[e_i'/x_i][v/x] \rhd (\Delta', s[\mathtt{p}_1] : \mathcal{L}_j)[v/x] \quad (4)$$

By applying $\lfloor$SEL$\rfloor$ with premises (3) and (5) we obtain the thesis. If $x : \langle\mathcal{L}\rangle$ the case is similar except x does not need to be substituted to the predicates.

Case [BCH]. We set $P = s[\mathtt{p},\mathtt{q}]?\{l_i(x_i)\langle E_i\rangle.P_i\}_{i\in I}$ and $\Delta = \Delta', s[\mathtt{p}_2] : \mathtt{p}_1?\{l_i(x_i : U_i)\{A_i\}\langle E_i\rangle.\mathcal{L}_j\}_{i\in I}$. By rule $\lfloor$BCH$\rfloor$ (we omit the case for delegation acceptance as it is similar) and assume $x : S'$:

$$\frac{\mathcal{C} \wedge A_i; \Gamma \vdash P_i \triangleright \Delta, s[\mathtt{p}_2] : \mathcal{L}_i \qquad \mathcal{C} \wedge A_i \supset \mathcal{C}\,\mathtt{after}\,E_i}{\mathcal{C}; \Gamma \vdash P \triangleright \Delta', s[\mathtt{p}_2] : \mathtt{p}_1?\{l_i(x_i : U_i)\{A_i\}\langle E_i\rangle.\mathcal{L}_j\}_{i\in I}}$$

Without loss of generality we assume $x \notin \{x_i\}_{i\in I}$. By induction

$$(\mathcal{C} \wedge A_i)[v/x]; \Gamma \vdash P_i[v/x] \triangleright (\Delta', s[\mathtt{p}_2] : \mathtt{p}_1?\{l_i(x_i : S_i)\{A_i\}\langle E_i\rangle.\mathcal{L}_j\}_{i\in I})[v/x] \quad (5)$$

Furthermore (proceeding as in (3)):

$$\mathcal{C} \wedge A_i[v/x] \supset \mathcal{C}\,\mathtt{after}\,E_i[v/x] \quad (6)$$

By applying (6) and (5) as a premise for $\lfloor$BCH$\rfloor$ we obtain the thesis.

Case [MREQ] *(resp.* [MACC]*).* This case follows straightforwardly by induction. The case for [MACC] is similar.

Case [VAR]. We set $P = X\langle e\rangle$. By $\lfloor$VAR$\rfloor$

$$\frac{\mathcal{L}_1[e/y]...\mathcal{L}_n[e/y] \text{ well-asserted}}{\mathcal{C}; \Gamma, X : (y : S')\mathcal{L}_1 @ \mathtt{p}_1..\mathcal{L}_n @ \mathtt{p}_n \vdash X\langle e\rangle \triangleright \Delta', s[\mathtt{p}_1] : \mathcal{L}_1[e/y], .., s[\mathtt{p}_n] : \mathcal{L}_n[e/y]}$$

Without loss of generality we assume $x \neq y$. Since $\mathcal{L}_1[e/y]...\mathcal{L}_n[e/y]$ are well-typed wrt the underlying typing discipline, $x : S, y : S'$ and $v : S$ then $\mathcal{L}_1[e/y][v/x]...\mathcal{L}_n[e/y][v/x]$ are also well-typed. $\mathcal{L}_1[e/y][v/x]...\mathcal{L}_n[e/y][v/x]$ are well-asserted by Lemma 16. By applying $\mathcal{L}_1[e/y][v/x]...\mathcal{L}_n[e/y][v/x]$ as a premise of $\lfloor$VAR$\rfloor$ we obtain the thesis.

Remaining Cases The other case are straightforward.

Evaluation

Lemma 2 (Evaluation). *If $\mathcal{C}; \Gamma \vdash P(e) \triangleright \Delta(e)$ and $\sigma \models e \downarrow v$ for a σ s.t. $\sigma \models \mathcal{C}$ then we have $\mathcal{C}; \Gamma \vdash P[e/v] \triangleright \Delta[e/v]$.*

Proof. The proof is by induction on the validation rules. We proceed by case analysis. By decidability of underlying logic, we can write $\sigma \models A[e/x] \downarrow \mathsf{true}$ when a closed formula $A[e/x]$ evaluates to true. Note that if we further have $e \downarrow$ then we have $A[v/x] \downarrow \mathsf{true}$.

Case [SEL]. If $P(e) = s[\mathtt{p},\mathtt{q}]!\{e_i \mapsto l_i\langle e_i'\rangle(x_j)\langle E_j\rangle; P_j(e)\}_{i\in J}$ then $P(v) = s[\mathtt{p},\mathtt{q}]!\{e_i \mapsto l_i\langle e_i'\rangle(x_j)\langle E_j\rangle; P_j(v)\}_{i\in J}$ and

$$\Delta(e) = \Delta = \Delta'(e), s[\mathtt{p}_1] : \mathtt{p}_2!\{l_i(x_i : S_i)\{A_i(e)\}\langle E_i(e)\rangle.\mathcal{L}_j(e)\}_{i\in I}$$

with and $\mathcal{C} \wedge e_i \supset (A_j \wedge (E_i = E_j) \wedge \mathcal{C}\,\mathtt{after}\,E_i)[e/x_j]$. Notice that $\mathcal{C} \wedge e_i \supset (A_j \wedge (E_i = E_j) \wedge \mathcal{C}\,\mathtt{after}\,E_i)[e/x_j]$ is equivalent to

$$\mathcal{C} \wedge e_i \supset (A_j \wedge (E_i = E_j) \wedge \mathcal{C}\,\mathtt{after}\,E_i)[v/x_j] \tag{7}$$

By inductive hypothesis

$$\mathcal{C}; \Gamma \vdash P'[v/e] \rhd \Delta'[v/e], s[\mathtt{p}] : \mathcal{L}[v/e] \tag{8}$$

By applying (7) and (8) to the validation rule [SEL] the lemma holds for this case.

Recursion Invocation If $P(e) = X\langle e\rangle$ (since $P(e)$ is well-formed against Δ by hypothesis) then $P(v) = X\langle v\rangle$. Since the substituted specification is still well-asserted (as it does not contain expressions) then $P(v)$ is well-formed against $\Delta[v/x]$ by rule [VAR].

Remaining cases The remaining cases are similar to the previous ones or straightforward by induction.

Other lemmas

Lemma 3 (Assertion Reduction and Coherence). *If Δ is coherent and $\langle \Gamma; [\Delta]\sigma\rangle \xrightarrow{\tau} \langle \Gamma'; [\Delta']\sigma'\rangle$ then Δ' is again coherent.*

Lemma 4 (Subject Congruence). *If $\mathcal{C}; \Gamma \vdash P_1 \rhd \Delta$ and $P_1 \equiv P_2$ then $\mathcal{C}; \Gamma \vdash P_2 \rhd \Delta$*

F.2 Soundness Proof

Theorem 13 (Soundness for Initial Networks) follows immediately from Lemma 6 (Soundness for open Networks), via Lemma 5. Lemma 5 shows there is a conditional simulation between the closing substitution of each single open validated located process and its corresponding specification. Recall that in the derivation of an open located process $\mathcal{C}$ may not be true, and we take a closing substitution consistent with $\mathcal{C}$ and Γ).

Lemma 5. *Let $\mathcal{R}$ be a relation collecting all pairs of the form $([P\tilde{\sigma}]\sigma_p; \langle \Gamma\tilde{\sigma}; [\Delta\tilde{\sigma}]\sigma_a\rangle)$ such that $\mathcal{C}; \Gamma \vdash P \rhd \Delta$ where:*

1. *$[P]\sigma_p$ is a sub-term of a multi-step $\xrightarrow{\ell}$-derivative of a located process,*
2. *$[\Delta]\sigma_a$ is an assertion assignment with state,*
3. *$(\_)\tilde{\sigma}$ is a closing substitution of interaction variables consistent with $\mathcal{C}$, Γ and Δ,*
4. *$\sigma_p \models \mathcal{I}(\sigma_p)$ and $\sigma_a \models \mathcal{I}(\sigma_p)$*

Then $\mathcal{R}$ is a conditional stateful simulation.

Proof. We show that $\mathcal{R}$ is a conditional stateful simulation by induction on the depth of the validation tree. We proceed by case analysis of the last rule applied.

Case [MREQ] *(resp.* [MACC]*).* In this case P is defined as $\overline{a}[\mathtt{n}](y).P'$. The last derivation rule for P is [MREQ] where $\Gamma = \Gamma', a : \mathcal{G}$

$$\frac{\mathcal{C}; \Gamma \vdash P' \rhd y[\mathtt{1}] : \mathcal{G} \upharpoonright \mathtt{1}, \Delta}{\mathcal{C}; \Gamma \vdash \overline{a}[\mathtt{n}](y).P' \rhd \Delta} \tag{9}$$

The only possible transition of $P\tilde{\sigma} = \overline{a}[\mathtt{n}](y).P'\tilde{\sigma}$ is by [TR-MREQ] in Figure 3. Notice that by $\sigma_a \models \mathcal{I}(\sigma_p)$ (condition (4) in the hypothesis).

By [TR-MREQ]:

$$[\overline{a}[\mathtt{n}](y).P'\tilde{\sigma}]_{\sigma_p} \xrightarrow{\overline{a}[\mathtt{n}]\langle s \rangle} [P'\tilde{\sigma}[s/y]]_{\sigma_p}$$

$\langle \Gamma\tilde{\sigma}; [\Delta\tilde{\sigma}]\sigma_a \rangle$ can move by [TR-A-MREQ] in Figure 9:

$$\langle \Gamma\tilde{\sigma}; [\Delta\tilde{\sigma}]\sigma_a \rangle \xrightarrow{\overline{a}[n]\langle s \rangle} \langle \Gamma\tilde{\sigma}; [\Delta\tilde{\sigma}, s[\mathtt{1}] : \mathcal{L}\tilde{\sigma}]\sigma_a \rangle$$

$([P'\tilde{\sigma}]\sigma_p; \langle \Gamma\tilde{\sigma}; [\Delta\tilde{\sigma}, s[\mathtt{1}] : \mathcal{L}\tilde{\sigma}]\sigma_a \rangle) \in R$ by applying Lemma 1 to the premise of (9), observing that the conditions (1÷4) are preserved. $\mathcal{R}$ is a conditional stateful simulation by induction. The case for [MACC] is similar.

Case [BCH]. In this case P is defined as $s?\{l_i(x_i)\langle E_i \rangle.P_i\}_{i \in I}$. The last derivation rule for P is [BCH]:

$$\frac{\begin{array}{c}\forall i \in I \quad \mathcal{C} \wedge A_i; \Gamma \vdash P_i \rhd \Delta' \qquad \mathcal{C} \wedge A_i \supset \mathcal{C}\,\mathtt{after}\,E_i \\ \text{if } U_i = \langle \mathcal{L} \rangle \text{ then } \Delta' = \Delta, s[\mathtt{q}] : \mathcal{L}_i, x_i : \mathcal{L} \text{ otw } \Delta' = \Delta, s[\mathtt{q}] : \mathcal{L}_i\end{array}}{\mathcal{C}; \Gamma \vdash s[\mathtt{p}, \mathtt{q}]?\{l_i(x_i)\langle E_i \rangle.P_i\}_{i \in I} \rhd \Delta', s[\mathtt{q}] : \mathtt{p}?\{l_i(x_i : U_i)\{A_i\}\langle E_i \rangle.\mathcal{L}_i\}_{i \in I}} \tag{10}$$

The possible transitions of $P\tilde{\sigma}$ are with label $s[\mathtt{p}, \mathtt{q}]?l_j\langle v \rangle$ for some v by [TR-BCH] in Figure 3:

$$[s[\mathtt{p}, \mathtt{q}]?\{l_i(x_i)\langle E_i\tilde{\sigma} \rangle.P_i\tilde{\sigma}\}_{i \in I}]\sigma_p \xrightarrow{s[\mathtt{p},\mathtt{q}]?l_j\langle v \rangle} [P_j\tilde{\sigma}[v/x_j]]\sigma'_p \quad \sigma'_p = \sigma_p\,\mathtt{after}\,E_j$$

By the shape of the specification in (10), $s[\mathtt{q}] : \mathtt{p}?\{l_i(x_i : U_i)\{A_i\}\langle E_i \rangle.\mathcal{L}_i\}_{i \in I}$ is able to make a move at subject $s?$. By definition of conditional simulation we are interested in inspecting only the case in which the specification can make a step with label $s[\mathtt{p}, \mathtt{q}]?l_j\langle v \rangle$. If the specification moves with label $s[\mathtt{p}, \mathtt{q}]?l_j\langle v \rangle$ we have two cases:

– Case $U_j = S$. Hence, by [TR-A-BCH] in Figure 9:

$$\langle \Gamma\tilde{\sigma}; [s[\mathtt{q}] : \mathtt{p}?\{l_i(x_i : U_i)\{A_i\tilde{\sigma}\}\langle E_i\tilde{\sigma} \rangle.\mathcal{L}_i\tilde{\sigma}\}_{i \in I}]\sigma_a \rangle \xrightarrow{s[\mathtt{p},\mathtt{q}]?l_j\langle v \rangle} \langle \Gamma\tilde{\sigma}; [s[\mathtt{q}] : \mathcal{L}_j\tilde{\sigma}]\sigma'_a \rangle$$

with

$$\sigma'_a = \sigma_a\,\mathtt{after}\,E_j \tag{11}$$

$([P_j\tilde{\sigma}]\sigma'_p; \langle \Gamma\tilde{\sigma}; \langle \Delta\tilde{\sigma}, s[\mathtt{q}] : \mathcal{L}_j\tilde{\sigma} \rangle \sigma'_a \rangle) \in R$ holds observing that the conditions (1÷4) are preserved. Notice that condition (4) follows by the second premise of (10). $\mathcal{R}$ is a conditional stateful simulation by induction.

– Case $U_j = \langle \mathcal{L} \rangle$. In this case $v = t[\mathtt{r}]$. This case is similar to the previous one except Δ moves by [TR-A-DELBCH] in Figure 9:

$$\begin{array}{l}\langle \Gamma\tilde{\sigma}; [s[\mathtt{q}] : \mathtt{p}?\{l_i(x_i : U_i)\{A_i\tilde{\sigma}\}\langle E_i\tilde{\sigma}\rangle.\mathcal{L}_i\tilde{\sigma}\}_{i\in I}]\sigma_a\rangle \xrightarrow{s[\mathtt{p},\mathtt{q}]?l_j\langle t[\mathtt{r}]\rangle} \\ \langle \Gamma\tilde{\sigma}; [s[\mathtt{q}] : \mathcal{L}_j\tilde{\sigma}, t[\mathtt{r}] : \mathcal{L}]\sigma'_a\rangle\end{array}$$

with

$$\sigma'_a = \sigma_a \,\mathtt{after}\, E_j \tag{12}$$

Case [SEL]. In this case P is defined as $s[\mathtt{p},\mathtt{q}]!\{e_i \mapsto l_i\langle e'_i\rangle(x_i)\langle E_i\rangle; P_i\}_{i\in I}$. The last derivation rule for P is [SEL]:

$$\frac{\begin{array}{c}\forall i \in I \exists j \in J \text{ s.t. } x_i = x_j \quad l_i = l_j \quad \mathcal{C} \wedge e_i \supset (A_j \wedge (E_i = E_j) \wedge \mathcal{C}\,\mathtt{after}\, A_j)[e'_i/x_i] \\ \mathcal{C} \wedge \underline{e_i}; \Gamma \vdash P[e'_i/x_i] \rhd \Delta', s[\mathtt{p}] : \mathcal{L}_j[e'_i/x_i] \\ \text{if } U_i = \langle \mathcal{L} \rangle \text{ then } \Delta = \Delta'', e'_i : \mathcal{L}'_i \text{ and } \Delta' = \Delta'' \text{ otw } \Delta' = \Delta\end{array}}{\mathcal{C}; \Gamma \vdash s[\mathtt{p},\mathtt{q}]!\{e_i \mapsto l_i\langle e'_i\rangle(x_i)\langle E_i\rangle; P_i\}_{i\in I} \rhd \Delta, s[\mathtt{p}] : \mathtt{q}!\{l_j(x_j : U_j)\{A_j\}\langle E_j\rangle.\mathcal{L}_j\}_{j\in J}} \tag{13}$$

The only possible transition is by [TR-SEL] in Figure 3. Assume first that the value sent is of type S. By [TR-SEL] then $P\tilde{\sigma}$ performs the following transition

$$[s[\mathtt{p},\mathtt{q}]!l_i(x_i)\langle E_i\tilde{\sigma}\rangle.P_i\tilde{\sigma}]_{\sigma_p} \xrightarrow{s[\mathtt{p},\mathtt{q}]!l_i\langle v\rangle} [P_i\tilde{\sigma}[v/x_i]]_{\sigma'_p} \quad \sigma'_p = \sigma_p \,\mathtt{after}\, E_i$$

Notice that

$$\sigma_a \models A_j\tilde{\sigma}[v/x_i] \tag{14}$$

following by $x_i = x_j, \sigma_a \models \mathcal{I}(\sigma_p)$ (condition (4) in the hypothesis), $\mathcal{C} \wedge e_i \supset A_j[v/x_i]$ (second premise of (13) above), and the fact that $\_\tilde{\sigma}$ is consistent with $\mathcal{C}$ (condition (3) in the hypothesis).

By (14) as premise of rule [TR-A-SEL] in Figure 9:

$$\langle \Gamma\tilde{\sigma}; [\Delta\tilde{\sigma}, s[\mathtt{p}] : \mathtt{q}!\{l_j(x_j : U_j)\{A_j\tilde{\sigma}\}\langle E_j\tilde{\sigma}\rangle.\mathcal{L}_j\tilde{\sigma}\}_{j\in J}]\sigma_a\rangle \xrightarrow{s[\mathtt{p},\mathtt{q}]!l_j\langle v\rangle} \langle \Gamma\tilde{\sigma}; [\Delta\tilde{\sigma}, s[\mathtt{p}] : \mathcal{L}_j]\sigma'_a\rangle$$

with $l_i = l_j$ and

$$\sigma'_a = \sigma_a \,\mathtt{after}\, E_j\tilde{\sigma} \tag{15}$$

$([P_i\tilde{\sigma}]_{\sigma_p}; \langle \Gamma\tilde{\sigma}; [\Delta\tilde{\sigma}, s[\mathtt{p}] : \mathcal{L}_j\tilde{\sigma}]\sigma'_a\rangle) \in R$, observing that the conditions $(1 \div 4)$ are preserved. Notice that case (4) holds by invariant stability. $\mathcal{R}$ is a conditional stateful simulation by induction.

This case in which x_j is a session channel is similar to the previous one, except transition [TR-A-SEL] of P has a corresponding [TR-A-DELSEL] of $\langle \Gamma\tilde{\sigma}; [\Delta\tilde{\sigma}]\sigma_a\rangle$.

Case [EMPTY]. We can set $P = \mathbf{0}$; the property holds since there are no transitions.

Case [PAR]. A parallel process $P = P_1 \mid P_2$ can only make either independent actions or reductions involving only either P_1 or P_2 (no reductions due to communication between P_1 and P_2 as only one role can be played by one principal in each session instance). This case is direct from the induction hypothesis.

Case [VAR]. We set P to be $X\langle e\rangle$ with $\Gamma(X) = (x : S)\mathcal{L}_1 @ \mathtt{p}_1..\mathcal{L}_n @ \mathtt{p}_n$. $P\tilde{\sigma}$ is a process such that $\mathcal{C}\tilde{\sigma}; \Gamma\tilde{\sigma} \vdash P\tilde{\sigma}[e/x] \triangleright \Delta\tilde{\sigma}$ where $\Delta\tilde{\sigma} = \Delta_0\tilde{\sigma}, s[\mathtt{1}] : \mathcal{L}_1[e/x], ..., s[\mathtt{n}] : \mathcal{L}_n[e/x]$ is the closure of the endpoint assertion of P. The property follows from the cases for the other process types.

Cases [REC]. This case is proved by the standard syntactic approximation of a recursion. We can assume, in all derivations for processes in P, the application of [REC] only occurs in (the last steps of) a derivation. Assume that we have

$$\mathcal{C}; \Gamma, X : (x : S)\mathcal{L}_1 @ \mathtt{p}_1..\mathcal{L}_n @ \mathtt{p}_n \models P \triangleright s[\mathtt{p}_1] : \mathcal{L}_1..s[\mathtt{p}_n] : \mathcal{L}_n \tag{16}$$

Further we also assume

$$\mathcal{C}; \Gamma, X : (x : S)\mathcal{L}_1 @ \mathtt{p}_1..\mathcal{L}_n @ \mathtt{p}_n \models Q \triangleright \Delta \tag{17}$$

Let y range over interaction names and session channels. In the following we often use the notation for the substitution $Q[(y)R/X]$ which replaces each occurrence of $X\langle e\rangle$ with $R[e/y]$. Using well-guardedness of process variables in [21], we first approximate the recursion by the following hierarchy:

$$P^0 \stackrel{def}{=} P' \sim \mathbf{0} \quad P^1 \stackrel{def}{=} P[(x)P^0/X] \quad \ldots \quad P^{n+1} \stackrel{def}{=} P[(x)P^n/X]$$

Above P^0 is chosen as the process which is typed by the same typing as P and which has no visible action. For example, choosing a and s to be fresh, $P^0 \stackrel{def}{=} (\nu a : \mathcal{G})(a[\mathtt{2}](y).P')$ then $P^0 \sim \mathbf{0}$. We also set $P^\omega = \mu X\langle y\rangle(x).P'$ to be the recursively defined agent itself.

In the conclusion of [REC] we abstract the process variable X by the μ construct. Instead, we replace each X in Q with $(y)P^0, (y)P^1, \ldots, (y)P^n$, and finally $(y)P^\omega$. We call the result $Q^0, Q^1, \ldots Q^n$, and Q^ω, where Q^ω is nothing but the term in the conclusion (after one-time unfolding which does not change the behaviour).

Now suppose that $\mathcal{C}; \Gamma \vdash S \triangleright \Delta$ is derivable and that $\mathcal{C}_0; \Gamma_0 \vdash S_0 \triangleright \Delta_0$ occurs in its derivation, hence $S0$ occurs in S. Suppose that also $\mathcal{C}_0; \Gamma_0 \vdash S_0' \triangleright \Delta_0$. We can replace the occurrence of S_0 in S by S_0', with the result written S', such that $\mathcal{C}; \Gamma \vdash S' \triangleright \Delta$ is derivable.

Using property, we first note that, for any $\langle \Gamma; [\Delta]\sigma_a\rangle$ and $\mathcal{C}$, we have $\mathcal{C}; \Gamma \models P^0 \triangleright \Delta$. Thus we apply this to (16) and replace X in P by $(x)P^0$:

$$\mathcal{C}; \Gamma \models P^1 \triangleright s[\mathtt{p}_1] : \mathcal{L}_1, \ldots, s[\mathtt{p}_n] : \mathcal{L}_n$$

This can again be used for (16) (noting the environment Γ can always be taken as widely as possible in [VAR]): $\mathcal{C}; \Gamma \models P^2 \triangleright s[\mathtt{p}_1] : \mathcal{L}_1, \ldots, s[\mathtt{p}_n] : \mathcal{L}_n$. In this way we know that for an arbitrary n: $\mathcal{C}; \Gamma \models P^n \triangleright s[\mathtt{p}_1] : \mathcal{L}_1, \ldots, s[\mathtt{p}_n] : \mathcal{L}_n$.

By applying this to (16) , we obtain:

$$\mathcal{C}; \Gamma \models Q^n \triangleright \Delta$$

for an arbitrary n. Now assume, for simplicity, that there are no free variables in Q (hence in Q^n) and therefore $\mathcal{C} = \mathsf{true}$ (the reasoning is precisely the same by applying a closing substitution). We can then construct a relation taking each node in the transitions from Q^ω and relating it to the derivative of $\langle \Gamma; [\Delta]\sigma_a \rangle$, by observing that assertions transitions are always deterministic for the given process and its transition derivatives. Let the resulting relation be $\mathcal{R}$. Since any finite trace of Q^ω is in some Q^n, the conditions of conditional simulation hold at each step.

Lemma 6 (Soundness for Open Networks). *Let N be a network. Then $\mathcal{C}; \Gamma \vdash N \rhd \Sigma$ implies $\mathcal{C}; \Gamma \models N \rhd \Sigma$*

Proof. Let $\mathcal{R}$ be a relation collecting all pairs of the form $(N; \Sigma)$ such that $\mathcal{C}; \Gamma \vdash N \rhd \Sigma$ where: (i) N is a sub-term of a multi-step $\xrightarrow{\ell}$-derivative of an initial network, (ii) Σ is a specification. Proceeding by induction on the length of the derivation tree. We proceed by case analysis of the validation rules for networks in Figure 4 and in Figure 11. Subject reduction for silent actions (Lemma 9)

Case [N1]. If [N1] is applied then $N = [P_i]\sigma_a$, $\Sigma = [\Delta']\sigma_p$, and

$$\mathcal{C}; \Gamma \vdash P \rhd \Delta$$

This case follows by Lemma 5, observing that by premise fourth condition ($\sigma_a \models \mathcal{I}(\sigma_p)$ and $\sigma_p \models \mathcal{I}(\sigma_p)$) holds by premise of [N1].

Case [N2]. This case follows by definition of refinement.

Case [N3]. This case is immediate since $N = \varnothing$ thus cannot make any transition.

Case [N4]. A parallel network $N = N_1 \mid N_2$ can make either independent actions or reductions. The case for independent actions is direct from the induction hypothesis. If the reduction takes place by interaction, then we use Lemma 9.

Cases [QNIL], [QVAL]. Queues do not have transitions. The behaviours of queues are taken into account as part of τ-actions in the case for [N3] above.

Case [CRES]. This case follows by Lemma 3.

G Subject Reduction Proofs

Lemma 7. *If $N \longrightarrow N'$ then one of the following cases hold:*

1. $N \equiv \mathcal{E}[\prod_{i\in\{1..n\}}[P_i]\sigma_i]$ *with* $P_1 = \overline{a}[\mathtt{n}](y_1).P_1' \mid Q_1$ *and* $P_i = a[\mathtt{i}](y_i).P_i' \mid Q_i$ *s.t.*
$$[P_1]\sigma_1 \xrightarrow{\overline{a}[\mathtt{n}]\langle s\rangle} [P_1' \mid Q_1]\sigma_1 \qquad [P_i]\sigma_i \xrightarrow{\overline{a}[\mathtt{i}]\langle s\rangle} [P_i' \mid Q_i]\sigma_i$$
and $N' = \mathcal{E}[(\nu s)(s : \varnothing \mid \prod_{i\in\{1..n\}}[P_i' \mid Q_i]\sigma_i)]$

2. $N \equiv \mathcal{E}[[P]\sigma \mid s : h]$ *s.t.* $[P]\sigma \xrightarrow{s[\mathtt{p},\mathtt{q}]!l\langle v\rangle} [P']\sigma$ *and* $N' \equiv \mathcal{E}[[P']\sigma \mid s : h \cdot (\mathtt{p}, \mathtt{q}, l\langle v\rangle)]$
3. $N \equiv \mathcal{E}[[P]\sigma \mid s : (\mathtt{p}, \mathtt{q}, l\langle v\rangle) \cdot h]$ *s.t.* $[P]\sigma \xrightarrow{s[\mathtt{p},\mathtt{q}]?l\langle v\rangle} [P']\sigma$ *and* $N' \equiv \mathcal{E}[[P']\sigma \mid s : h]$

Proof. Immediate from the corresponding reduction rules.

Lemma 8. *1. If* $[P]\sigma \xrightarrow{a[\mathtt{n}]\langle s\rangle} [P']\sigma$ *and* $\mathcal{C}; \Gamma \vdash [P]\sigma \rhd \Sigma$ *then* $\Sigma = [\Delta]\sigma'$ *for some* Δ, σ', *and* $\mathcal{C}; \Gamma \vdash [P']\sigma \rhd [\Delta, s[\mathtt{1}] : \mathcal{L}]\sigma'$

2. *If* $[P]\sigma \xrightarrow{a[\mathtt{i}]\langle s\rangle} [P']\sigma$ *and* $\mathcal{C}; \Gamma \vdash [P]\sigma \rhd \Sigma$ *then* $\Sigma = [\Delta]\sigma'$ *for some* Δ, σ', *and* $\mathcal{C}; \Gamma \vdash [P']\sigma \rhd [\Delta, s[\mathtt{i}] : \mathcal{L}]\sigma'$

3. *If* $[P]\sigma_p \xrightarrow{s[\mathtt{p},\mathtt{q}]!l\langle v\rangle} [P']\sigma'_p$ *and* $\mathcal{C}; \Gamma \vdash [P]\sigma_p \mid s : h \rhd \Sigma$ *then* $\Sigma = [\Delta]\sigma_a$ *for some* Δ, σ_a *and* $\mathcal{C}; \Gamma \vdash [P']\sigma_p \mid s : h \cdot (\mathtt{p}, \mathtt{q}, l\langle v\rangle) \rhd [\Delta']\sigma'_a$ *s.t.* $\langle \Gamma; [\Delta]\sigma_a\rangle \xrightarrow{\tau} \langle \Gamma; [\Delta']\sigma'_a\rangle$

4. *If* $[P]\sigma_p \xrightarrow{s[\mathtt{p},\mathtt{q}]?l\langle v\rangle} [P']\sigma'_p$ *and* $\mathcal{C}; \Gamma \vdash [P]\sigma \mid s : (\mathtt{p}, \mathtt{q}, l\langle v\rangle) \cdot h \rhd \Sigma$ *then* $\Sigma = [\Delta]\sigma'$ *for some* Δ, σ', *and either*
   - Σ *can move at subject* $s?[\mathtt{p}, \mathtt{q}]$ *but cannot move with label* $s[\mathtt{p}, \mathtt{q}]?l\langle v\rangle$
   - $\mathcal{C}; \Gamma \vdash [P']\sigma_p \mid s : h \rhd [\Delta']\sigma_a$ *s.t.* $\langle \Gamma; [\Delta]\sigma_a\rangle \xrightarrow{\tau} \langle \Gamma; [\Delta']\sigma'_a\rangle$.

Proof. (1) and (2) are immediate. Below we show the cases (3) and (4).

Case (3) Suppose we have $\mathcal{I}(\sigma_p); \Gamma \vdash P \mid s : h \rhd \Delta$. We safely assume the last rule applied is [PAR], thus we can assume $\Delta = \Delta_0, \Delta_1$ for some Δ_0 and Δ_1, and

$$\mathcal{I}(\sigma_p); \mathcal{C}; \Gamma \vdash P \rhd \Delta_0 \tag{18}$$

Now consider the transition $[P]\sigma_p \xrightarrow{s[\mathtt{p},\mathtt{q}]!l_i\langle v\rangle} [P']\sigma'_p$, by (18) we observe Δ_0 has the shape

$$\Delta_0 = s[\mathtt{p}] : \mathtt{q}!\{l_j(x_j : U_j)\langle E_j\rangle\{A_j\}; \mathcal{L}_j\}_{j\in J}\}, \Delta_{00}$$

and that P' can be typed by Δ'_0 such that:

$$\Delta'_0 = s[\mathtt{p}] : \mathcal{L}_j[v/x_i], \Delta_{00} \tag{19}$$

Now the assertion Δ_1 for the queue has the shape, omitting the vacuous "end": $\Delta_1 = s : \mathcal{M}$ hence the addition of the values to this queue, $s : h \cdot (\mathtt{p}, \mathtt{q}, l_i\langle v\rangle)$, must have the endpoint assertion:

$$\Delta'_1 = s[\mathtt{p}] : \mathtt{q}!l_i\langle v\rangle; \mathcal{M} \tag{20}$$

Setting $\Delta' = \Delta'_0, \Delta'_1$, we know $\mathcal{I}(\sigma_p); \Gamma \vdash P' \mid s : h \cdot (\mathtt{p}, \mathtt{q}, l_i\langle v\rangle) \rhd \Delta'$. By (19) and (20) we obtain

$$\Delta'_0, \Delta'_1 = s[\mathtt{p}] : \mathtt{q}!l_i\langle v\rangle; \mathcal{L}_j[v/x_i], \Delta_{00}, \Delta_1$$

and

$$\langle \Gamma; \Delta_0, \Delta_1\rangle \xrightarrow{\tau} \langle \Gamma; \Delta'_0, \Delta'_1\rangle$$

and the only change is at the assertion assignment at $s[\mathtt{p}]$, as required.

Case (4). Suppose we have $\mathcal{I}(\sigma_p); \Gamma \vdash P \mid s : h \cdot (\mathtt{p}, \mathtt{q}, l_i\langle v\rangle) \triangleright \Delta$. Again we safely assume the last rule applied is [PAR]. Thus we can assume, for some Δ_0 and Δ_1:

$$\mathcal{I}(\sigma_p); \Gamma \vdash s : h \cdot (\mathtt{p}, \mathtt{q}, l_i\langle v\rangle) \triangleright \Delta_1$$

with $\Delta = \Delta_0, \Delta_1$, and

$$\mathcal{I}(\sigma_p); \Gamma \vdash P \triangleright \Delta_0 \tag{21}$$

Now consider the transition

$$[P]\sigma_p \xrightarrow{s[\mathtt{p},\mathtt{q}]?l_j\langle v\rangle} [P']\sigma'_p \tag{22}$$

As before, we can infer, from (21) and (22) the shape of Δ_0 as follows,

$$\Delta_0 = s[\mathtt{q}] : \mathtt{p}?\{l_j(x_i : U_i)\langle E_i\rangle\{A_i\}; \mathcal{L}_i\}_{i\in I}, \Delta_{00}$$

for some $\mathtt{p}$; and that P' can be validated against Δ'_0 given as

$$\Delta'_0 = s[\mathtt{q}] : \mathcal{L}[v/x_j], \Delta_{00} \tag{23}$$

Now the assertion Δ_1 for the queue has the shape (again omitting "end"-only assertions):

$$\Delta_1 = s[\mathtt{q}] : \mathtt{p}?l_j\langle v\rangle; \mathcal{M} \tag{24}$$

which, if we take off the values (hence for the queue $s : h$), we obtain:

$$\Delta'_1 = s[\mathtt{q}] : \mathcal{M} \tag{25}$$

Note this is symmetric to the case (1) above. As before, setting $\Delta' = \Delta'_0, \Delta'_1$, we know: $\mathcal{I}; \mathcal{C}; \Gamma \vdash P' \mid s : h \triangleright \Delta'$. By (23) and (25) we obtain

$$\begin{array}{rl} \Delta'_0, \Delta'_1 = & s[\mathtt{q}] : \mathcal{L}_j[v/x_j],\ \Delta_{00},\ \Delta_1 \\ \overset{\tau}{\leftarrow} & \Delta_0, \Delta_1 \end{array}$$

The only change from Δ to Δ' is at the type assignment at $s[\mathtt{q}]$, as required.

For convenience of the case analysis we explicitly write $P \xrightarrow{\tau_s} P'$ if $P \xrightarrow{\tau} P'$ is derived by the reduction rules for free session channels.

Lemma 9 (Subject Reduction for Silent Actions). *Suppose $\Gamma \vdash N \triangleright \Sigma$.*

1. *if $N \xrightarrow{\tau} N'$ then $\Gamma \vdash N' \triangleright \Sigma$ again*
2. *if $N \xrightarrow{\tau_s} N'$ then there exists Σ' s.t. $\langle \Gamma, \Sigma\rangle \xrightarrow{\tau} \langle \Gamma, \Sigma'\rangle$ and $\Gamma \vdash N' \triangleright \Sigma'$.*

Proof. If $N \xrightarrow{\tau} N'$ then each of the cases of Lemma 7 are possible, we inspect them one by one.

Case (1): Session Initiation. By Lemma 7 (1) and (2) we set $N = [P_1]\sigma_1 \mid \prod_{2 \geqslant i \geqslant n}[P_i]\sigma_i$ where $P_1 = \overline{a}[\mathtt{n}](y_1).P_1' \mid Q_1$ and $Q_i = a[\mathtt{i}](y_i).P_i' \mid Q_i$. As given in Lemma 7 (2) the actions of P_i compensate each others and correspond to reduction $N \longrightarrow (\nu s)(s : \varnothing \mid \prod_{1 \geqslant n \geqslant n}[P_i'[s[\mathtt{i}]/y_i] \mid Q_i]\sigma_i)$ by the first rule in Figure 8.

Since $\mathcal{E}[\_]$ is a reduction context we can safely set

$$\Gamma \vdash [a[\mathtt{i}](y_i)\{A_i\}.P_i]\sigma_i \rhd [\Delta_i]\sigma_i'$$

so that $[\Delta_1]\sigma_1', ..., [\Delta_n]\sigma_n' = \Sigma$.

Hence, by premise of validation rule [MACC] we have, with $\Gamma(a) = \mathcal{G}$,

$$\Gamma \vdash [P_i]\sigma_i \rhd [\Delta_i, s[\mathtt{i}] : \mathcal{L}_i]\sigma_i'$$

with $\mathcal{G} \upharpoonright \mathtt{i} = \mathcal{L}_i$ (similarly for role 1).

Since $\{\mathcal{G} \upharpoonright \mathtt{i}\}_{i \in I}$ is obviously coherent then $\Gamma \vdash (\nu s)(\prod_i P_i'[s\mathtt{i}]/y] \mid Q_i) \rhd \Delta$ as required.

Case (2): Select. By Lemma 7 (3) we set $N \equiv \mathcal{E}[[P]\sigma_p \mid s : h]$ with $[P]_{\sigma_p} \xrightarrow{s[\mathtt{p},\mathtt{q}]!l\langle v\rangle} [P']_{\sigma_p'}$. As above we can safely set $\mathcal{I}(\sigma_p); \Gamma \vdash [Q]_{\sigma_p'} \mid s : \tilde{h} \rhd \Delta$. By Lemma 8 we can infer $\mathcal{I}(\sigma_p); \Gamma \vdash [P']_{\sigma_p'} \mid s : \tilde{h} \cdot (\mathtt{p}, \mathtt{q}, l\langle v\rangle) \rhd \Delta'$ such that $\langle \Gamma, [\Delta]\sigma_a\rangle \xrightarrow{\tau} \langle \Gamma, [\Delta']\sigma_a'\rangle$. Since N reduces to N' by τ-transition rather than τ_s transition, we know that s is hidden in N. Assume therefore without loss of generality

$$N \equiv C[(\nu s)([P]_\sigma \mid s : h \mid M)] \qquad \mathcal{I}(\sigma_p); \Gamma \vdash [P]_\sigma \mid s : h \mid M \rhd \Sigma_1$$

with Σ_1 coherent and $\Sigma_1 = \Sigma, \Sigma_{01}$. By Lemma 3 and $\Sigma_1 \longrightarrow \Sigma', \Sigma_{01}$ we know Σ', Σ_{01} is also coherent, hence done.

Case (3): Branch. The argument exactly follows case (2) above except using Lemma 7 (3) and Lemma 8 (4) instead of Lemma 7 (2) and Lemma 8 (3), respectively.

Case τ_s. Proceeds as above but without restricting s.

Lemma 10. *If $\mathcal{C}; \Gamma \vdash P \rhd \Delta$ and $[P]\sigma_p \xrightarrow{\ell} [P']\sigma_p'$ for some ℓ, P', σ_p and σ_p s.t. $\sigma_p \models \mathcal{I}(\sigma_p)$ then:*

- *if ℓ is a branching action then $\langle \Gamma; [\Delta]\sigma_p\rangle$ is able to move at subject of ℓ, and if $\langle \Gamma; [\Delta]\sigma_p\rangle \xrightarrow{\ell} \langle \Gamma; [\Delta']\sigma_a'\rangle$ then we have $\mathcal{I}(\sigma_p); \Gamma \vdash P' \rhd \Delta'$.*
- *if ℓ is not a branching nor a τ action then $\langle \Gamma; [\Delta]\sigma_a\rangle \xrightarrow{\ell} \langle \Gamma; [\Delta']\sigma_a'\rangle$ then we have $\mathcal{I}(\sigma_p); \Gamma \vdash P' \rhd \Delta'$.*

Proof. The proof is by induction on the validation rules .We proceed by the case analysis depending on the last rule used for deriving this judgement. We assume processes are closed. Further below notice $\mathcal{C}$ in the conclusion of each rule should be true by our assumption.

Case [SEL]. In this case, we derive $\mathcal{I}(\sigma_p); \Gamma \vdash P \rhd \Delta$ with:

$$P = s[\mathtt{p},\mathtt{q}]!\{e_i \mapsto l_i\langle e_i'\rangle(x_i)\langle E_i\rangle; P_i\}_{i\in j} \text{ and } \Delta = \Delta_0, s[\mathtt{p}] : \mathtt{q}!\{l_j(x_j : U_j)\{A_j\}\langle E_j\rangle; \mathcal{L}_j\}_{j\in J} \tag{26}$$

P can move only by [TR-SEL] in Figure 3: $[P]\sigma_p \xrightarrow{\ell} [P_i[v/x_i]]\sigma_p'$ with $\ell = s[\mathtt{p},\mathtt{q}]!l_i\langle v\rangle$, $\sigma_p' = \sigma_p \,\mathtt{after}\, E_i$, and

$$\sigma_p \models e_i \downarrow v \tag{27}$$

since P is closed the only free variables in e_i are state variables defined in σ_p. By the first premise of validation rule [SEL] we have:

$$\mathcal{I} \supset A_j[e_i'/x_j] \tag{28}$$

By (27) and (28) we infer $\mathcal{I} \supset A_j[v/x_j]$.

From $\mathcal{I} \supset A[v/x_j]$, since $\sigma_a \models \mathcal{I}$ we have $\sigma_a \models A[v/x_j]$ hence $\langle \Gamma; [\Delta]\sigma_a\rangle$ can move by [TR-SEL]:

$$\langle \Gamma; [\Delta]\sigma_a\rangle \xrightarrow{\ell} \langle \Gamma; [\Delta_0, s[\mathtt{p}] : \mathcal{L}_j[v/x_j]]\sigma_a'\rangle$$

with $\sigma_a' = \sigma_a \,\mathtt{after}\, E_i$

By the third premise of validation [SEL] we have

$$\mathcal{I}(\sigma_p); \Gamma \vdash P_j[e_i'/x_i] \rhd \Delta_0, s[\mathtt{p}] : \mathcal{L}_j[e_i'/x_j] \tag{29}$$

By Evaluation Lemma, (29) immediately gives $\mathcal{I}; \mathsf{true}; \Gamma \vdash P_j[v/x_j] \rhd \Delta_0, s[\mathtt{p}] : \mathcal{L}[v/x_j]$ as required. This case holds by induction observing that $\sigma_p' \models \mathcal{I}$ and $\sigma_a' \models \mathcal{I}$ by invariant stability.

Case [BCH]. In this case the conclusion is $\mathcal{I}(\sigma_p); \Gamma \vdash P \rhd \Delta$ with:

$$P = s[\mathtt{p},\mathtt{q}]?\{l_i(x_i)\langle E_i\rangle.P_i\}_{i\in I} \quad \text{and} \quad \Delta = \Delta_0, s[\mathtt{p}] : \mathtt{q}?\{l_i(x_i : U_i)\{A_i\}\langle E_i\rangle.\mathcal{L}_i\}_{i\in I} \tag{30}$$

By the shape of P we can set $\ell = s[\mathtt{q},\mathtt{p}]?l_j\langle v\rangle$. P can move only by [TR-BCH], obtaining $[P]\sigma_p \xrightarrow{\ell} [P_j[v/x_j]]\sigma_p'$, with $\sigma_p' = \sigma_p \,\mathtt{after}\, E_j$. By the shape of Δ from the validation rule [BCH] we have that Δ is able to move at subject of ℓ. In case $\langle \Gamma; [\Delta]\sigma_a\rangle$ can move with label ℓ we have by [TR-BCH]:

$$\langle \Gamma, [\Delta]\sigma_a\rangle \xrightarrow{\ell} \langle \Gamma; [\Delta_0, s[\mathtt{p}] : \mathcal{L}[v/x_j]]\sigma_a'\rangle$$

with $\sigma_a' = \sigma_a \,\mathtt{after}\, E_j$. for which $\sigma_a \models A[v/x_j]$. Now the premise of validation rule [BCH]:

$$\mathcal{I}(\sigma_p) \wedge A_j; \Gamma \vdash P_j \rhd \Delta_0, s[\mathtt{p}] : \mathcal{L}$$

By Substitution Lemma we obtain

$$\mathcal{I}(\sigma_p) \wedge A_j[v/x_j]; \Gamma \vdash P_j[v/x_j] \rhd \Delta_0, s[\mathtt{p}] : \mathcal{L}[v/x_j]$$

Since by history sensitivity A_j does not contain free state variables the it is possible to evaluate it. By $A_j[v/x_j] \downarrow \mathsf{true}$ and by validation rule [BCH] we obtain $\mathcal{I}(\sigma_p); \Gamma \vdash P_j[v/x_j] \rhd \Delta_0, s[\mathtt{p}] : \mathcal{L}[v/x_j]$ as required. This case holds by induction observing that $\sigma_p' \models \mathcal{I}$and $\sigma_a' \models \mathcal{I}$ by invariant stability.

Case [MREQ]. In this case we have $\mathcal{I}(\sigma_p); \Gamma \vdash P \triangleright \Delta$ such that, combining with the premises of the rule [MREQ] we have: $P = \overline{a}[n](y).Q$ and $\mathcal{I}(\sigma_p); \Gamma \vdash Q \triangleright \Delta, s[\mathtt{1}] : \mathcal{L}$ where

$$\Gamma(a) = \mathcal{G} \text{ and } \mathcal{G} \upharpoonright_1 = \delta\{A\}.\mathcal{L} \text{ and } \mathcal{I} \supset A \tag{31}$$

By the shape of P we can set $\ell = \overline{a}[\mathtt{n}]\langle s \rangle$ and $P \xrightarrow{\ell} Q$. By (31) we have $\mathcal{I} \supset A$ and by hypothesis $\sigma_a \models \mathcal{I}$ hence $\sigma_a \models A$. Therefore the following transition is possible using [TR-A-MREQ]: $\langle \Gamma, [\Delta]\sigma_p \rangle \xrightarrow{\ell} \langle \Gamma, [\Delta, s[\mathtt{1}] : \mathcal{L}]\sigma_a \rangle$ as required.

Case [MACC]. Similar to the case [MCAST] above.

Case [PAR]. Immediate, since the visible transition for $P \mid Q$ is reducible to the same action by either P or Q, and because the resulting assertion environments (one result of the visible transition) can again be composed, because linear compatibility only depends on channel names and participant names.

Case[REC]. This case follows from applying induction on the unfolding of P and folding it back after the transition.

Other cases. In each case, direct from the induction hypothesis.

Case [VAR]. Immediate since in this case there is no reduction from P.

Lemma 11. *If* $\mathsf{true}; \Gamma \vdash N \triangleright \Sigma$ *and* $N \xrightarrow{\ell} N'$ *and* $\langle \Gamma; \Sigma \rangle \xrightarrow{\ell} \langle \Gamma; \Sigma' \rangle$ *where* $\ell \neq \tau$*, then we have* $\mathsf{true}; \Gamma \vdash N' \triangleright \Sigma'$

Proof. The proof is by induction on the validation rules. The case for [N1] follows by Lemma 11. The cases for [N2] and [N3] are straightforward. We show below the case for [N4].

Suppose the conclusion is $\mathsf{true}; \Gamma \vdash N \triangleright [\Delta]\sigma$ which is derived from

$$\mathsf{true}; \Gamma_0 \vdash N \triangleright [\Delta_0]\sigma_0 \tag{32}$$

with $\Gamma_0, \Delta_0, \sigma_0 \ni \Gamma, \Delta, \sigma$. Now first suppose the concerned visible action ℓ is neither a receive action nor a branching. Now suppose $N \xrightarrow{\ell} N'$. By induction hypothesis and by (32), $\langle \Gamma_0; [\Delta_0]\sigma_0 \rangle \xrightarrow{\ell} \langle \Gamma_0'; [\Delta_0']\sigma_0' \rangle$ for some Γ_0', Δ_0' and σ_0' for which we have, by induction hypothesis

$$\mathsf{true}; \Gamma_0' \vdash N' \triangleright [\Delta_0']\sigma_0' \tag{33}$$

Since the assertion transition is deterministic and by definition of refinement $\Gamma_0', \Delta_0', \sigma_0' \ni \Gamma', \Delta', \sigma'$, by (33) we can use [N4] to reach the thesis.

H Completeness

We give here the outline proof for completeness. Assuming $\Gamma \models [P]\sigma_p \rhd [\Delta_0]\sigma_a$, we introduce generation rules to obtain a formula Δ parametric with respect to a number of predicate variables. Then we show that there exist a substitution ξ of the predicate variables in Δ such that: (1) $\mathsf{true}; \Gamma \vdash [P]\sigma_p \rhd [\Delta\xi]\sigma_a$ (i.e., provability of validation rules), and (2) $\Delta\xi \ni \Delta_0$ (completeness via refinement). The thesis is a consequence of (2) and of validation rule [N4]. The full proofs and related definitions are long, so that we post them in the link with the related materials [5].

I HML Embedding

Embedding We use a standard HML with the first-order predicates as in [1]. These predicates, denoted by A in the following are to the ones appearing in assertions, defined in Figure 1. The LTS associated to our HML consider as actions, denoted by ℓ, both the communications of the process and the updates of the state. As a consequence $P, \sigma \xrightarrow{\ell} P', \sigma'$ if either $P \xrightarrow{\ell} P'$ and $\sigma' = \sigma$ or $P = P'$ and $\sigma' = \sigma\,\mathtt{after}\,\ell$. We use ϕ to denote HML-formulae, which are built from predicates, implications, universal quantifiers, conjunctions and *must* modalities. We remark that the logic used in this *safety embedding* is positive: if we remove the implication symbol, there is no negation, no existential quantifier, no disjunction and no may modality. Additionally, the implication will always appear as $A \Rightarrow \phi$ meaning that modalities never appear in the negative side.

$$\phi ::= \mathtt{true} \mid \phi \wedge \phi \mid \phi \Rightarrow \phi \mid [\ell]\phi \mid A \mid \forall x : S.\phi \quad \ell ::= s[\mathtt{p}, \mathtt{q}](x) \mid \overline{s[\mathtt{p}, \mathtt{q}]}(x) \mid E$$

$$\frac{P, \sigma \models \phi_1 \qquad P, \sigma \models \phi_2}{P, \sigma \models \phi_1 \wedge \phi_2} \qquad \frac{}{P, \sigma \models \mathtt{true}}$$

$$\frac{\text{if } P, \sigma \models \phi_1 \text{ then } P, \sigma \models \phi_2}{P, \sigma \models \phi_1 \Rightarrow \phi_2} \qquad \frac{\text{For all } P', \sigma' \text{ s.t. } P, \sigma \xrightarrow{\ell} P', \sigma', P', \sigma' \models \phi}{P \models [\ell]\phi}$$

$$\frac{\sigma \vdash_{bool} A}{P, \sigma \models A} \qquad \frac{\text{For all values } v \text{ of type } T, P, \sigma \models \phi[v/x]}{P, \sigma \models \forall x : T.\phi}$$

Fig. 13. Logical rules

The satisfactions rules (Figure 14) are fairly standard, for a pair P, σ to satisfy a predicate A, A has to hold w.r.t. to σ, denoted by $\sigma \vdash_{bool} A$, meaning that $\sigma(A)$ is a tautology for the boolean logic.

$$\frac{P,\sigma \models \phi_1 \qquad P,\sigma \models \phi_2}{P,\sigma \models \phi_1 \wedge \phi_2} \qquad\qquad \frac{}{P,\sigma \models \texttt{true}}$$

$$\frac{\text{if } P,\sigma \models \phi_1 \text{ then } P,\sigma \models \phi_2}{P,\sigma \models \phi_1 \Rightarrow \phi_2} \qquad\qquad \frac{\text{For all } P',\sigma' \text{ s.t. } P,\sigma \xrightarrow{\ell} P',\sigma', P',\sigma' \models \phi}{P \models [\ell]\phi}$$

$$\frac{\sigma \vdash_{bool} A}{P,\sigma \models A} \qquad\qquad \frac{\text{For all values } v \text{ of type } T, P,\sigma \models \phi[v/x]}{P,\sigma \models \forall x : T.\phi}$$

Fig. 14. Logical rules

The embedding of local types we propose is parametrised with a session channel $s[\mathtt{p}]$. Assertions appearing in input prefixes are embedded as premises in implications, and assertions in output prefixes have to be satisfied, yielding:

$$\|\mathtt{q}!\{l_i(x_i : S_i)\{A_i\}\langle E_i\rangle.\mathcal{L}_i\}_{i\in I}\|^{s[\mathtt{p}]} = \textstyle\bigwedge_{i\in I} \forall x_i : S_i, [\overline{s[\mathtt{p},\mathtt{q}]}(x_i)](A_i \wedge [E_i]\|\mathcal{L}_i\|^{s[\mathtt{p}]})$$
$$\|\mathtt{q}?\{l_j(x_j : S_j)\{A_j\}\langle E_j\rangle.\mathcal{L}_j\}_{j\in J}\|^{s[\mathtt{p}]} = \textstyle\bigwedge_{j\in J} \forall x_j : S_j, [s[\mathtt{q},\mathtt{p}](x_j)](A_j \Rightarrow \|\mathcal{L}_j\|^{s[\mathtt{p}]})$$

The embedding of selection, is a conjunction of formulae corresponding to the branches: for each value sent on the session channel, predicates should be satisfied and, if the state is updated, the embedding of the continuation should hold. For branching types, the assertion is used as an hypothesis and no update appear.

Soundness To obtain soundness for typing judgements involving specifications, we have to introduce *interleavings* of formulae, treating the fact that one process can play several roles in several sessions. As a simple example both $s[\mathtt{p}_1,\mathtt{p}_2]?(x).k![\mathtt{q}_1,\mathtt{q}_2]\langle 10\rangle$ and $k![\mathtt{q}_1,\mathtt{q}_2]\langle 10\rangle.s[\mathtt{p}_1,\mathtt{p}_2]?(x)$ can be typed with $s[\mathtt{p}_2] : \mathtt{p}_1?(x : \mathtt{Nat}).\mathsf{end}$, $k[\mathtt{q}_1] : \mathtt{q}_2!(y : \mathtt{Nat}).\mathsf{end}$.

Interleaving is not a new operator *per se* and can be seen as syntactic sugar, describing shuffling of must modalities. The main rule for interleaving is: $[\ell_1]\phi_1 \rtimes [\ell_2]\phi_2 = [\ell_1](\phi_1 \rtimes [\ell_2]\phi_2) \wedge [\ell_2]([\ell_1]\phi_1 \wedge \phi_2)$. When interleaving two or more formulae containing modalities, we obtain a conjunction of formulae, each one representing a different way of organising all modalities in a way preserving their initial orders. Informally, the interleaving of $[1][2]$ and $[A][B]$ is $[1][2][A][B] \wedge [A][B][1][2] \wedge [1][A][2][B] \wedge [A][1][B][2] \wedge [1][A][B][2] \wedge [A][1][2][B]$.

We encode a pair Δ, Γ into a complex formula $\mathtt{Inter}(\Delta,\Gamma)$, defined as the interleaving of the formulae obtained by encoding the local types of Δ on their corresponding channels and the formulae corresponding to Γ, built as follows: for each channel $a : \mathtt{I}(\mathcal{G})$, if some $s[\mathtt{p}]$ is received on a, the resulting process should satisfy the encoding on $s[\mathtt{p}]$ of the projection of $\mathcal{G}$ on $\mathtt{p}$:

$$\mathtt{Inter}(s_1[\mathtt{p}_1],\ldots,s_n[\mathtt{p}_n]; a_1 : \mathtt{I}(\mathcal{G}_1),\ldots,a_m : \mathtt{I}(\mathcal{G}_m)) = \|T_1\|^{s_1[\mathtt{p}_1]} \rtimes \ldots \rtimes \|T_n\|^{s_n[\mathtt{p}_n]} \rtimes \phi_1 \rtimes \ldots \rtimes \phi_m$$

where $\phi_i = \forall s_i'.\forall \mathtt{p}_i'.[a_i(s_i'[\mathtt{p}_i])]\|\mathcal{G}_i \upharpoonright \mathtt{p}_i'\|^{s_i'[\mathtt{p}_i']}$.

The preliminaries lemmas concerning logics need to be proved. Lemma 12 states that a process cannot perform an action on a channel that does bot appear in its type. Lemma 13 observes that parallel composition with processes that does not perform any action does not change the set of formulae a process satisfies. Lemma 14 states that satisfaction of assertion is stable by reduction and Lemma 15 enforces the stability of satisfaction judgement by well-typed substitutions.

Lemma 12 (Type safety). *If $\mathcal{C}; \Gamma \vdash P : \Delta$ and $s[\mathtt{p}] \notin \Delta \cup \Gamma$, then there is no P' s.t. $P, \sigma \xrightarrow{\ell_s} P', \sigma$ for any action ℓ_s of the form $s![\mathtt{p}, \mathtt{q}]\langle l.v\rangle$ or $s![\mathtt{q}, \mathtt{p}]l(x)$.*

Similarly, if $a : \mathtt{I}(\mathcal{G}) \notin \Gamma$, there is no $P', s[\mathtt{p}]$ s.t. $P, \sigma \xrightarrow{a(s[\mathtt{p}])} P', \sigma$.

The direct corollary that will be used later is that a process typed with an empty Δ cannot make any action.

Lemma 13 (Trivial Composition). *If $P_1, \sigma \models \phi$ and P_2 cannot make any action, then $P_1 \mid P_2, \sigma \models \phi$.*

Lemma 14 (Stability of assertions). *If $P, \sigma \models A$ and $P \xrightarrow{\ell} P'$, then $P', \sigma \models A$.*

Lemma 15 (Satisfaction substitution). *If $P, \sigma \models \phi$ and $x : S, v : S$ are not bound in P, σ and ϕ, then $P[v/x], \sigma \models \phi[v/x]$.*

We state, thanks to the previous lemmas, the following 'simple' soundness, for simple local types:

Proposition 1 (Simple Soundness). *If $\mathcal{C}, \varnothing \vdash P \rhd s[\mathtt{p}] : \mathcal{L}$, then $(P, \sigma) \models \mathcal{C} \Rightarrow \|\mathcal{L}\|^{s[\mathtt{p}]}$.*

Unasserted types are built from:

$$\mathcal{L} ::= \quad \mathtt{p}?\{l_i(x_i : U_i)E_i.\mathcal{L}_i\}_{i\in I} \mid \mathtt{p}!\{l_i(x_i : S_i)E_i.\mathcal{L}_i\}_{i\in I} \mid \mathsf{end}$$

The multiplicative parallel rule is given as:

$$\frac{\mathcal{C}; \Gamma_1 \vdash P_1 : \Delta_1 \qquad \mathcal{C}; \Gamma_2 \vdash P_2 : \Delta_2}{\mathcal{C}; \Gamma_1, \Gamma_2 \vdash P_1 \mid P_2 : \Delta_1, \Delta_2}$$

Proposition 2 (Simple Completeness). *For all $\mathcal{L}$, if $\vdash P : s[p] : \mathbf{Er}(\mathcal{L})$ and $P, \sigma \models \mathcal{C} \Rightarrow \|\mathcal{L}\|^{s[\mathtt{p}]}$ then $\mathcal{C}; \vdash P \rhd s[p] : \mathcal{L}$.*

Here are additional definitions for interleaving:

$$[\ell_1]\phi_1 \rtimes (\phi_{2,1} \wedge \phi_{2,2}) = [\ell_1](\phi_1 \rtimes \phi_{2,1}) \wedge [\ell_1](\phi_1 \rtimes \phi_{2,2})$$

$$\phi \rtimes \mathtt{true} = \phi \qquad\qquad \phi \rtimes (\phi_1 \wedge \phi_2) = (\phi \rtimes \phi_1) \wedge (\phi \rtimes \phi_2)$$

$$(\phi_1 \wedge \phi_2) \rtimes \phi = (\phi_1 \rtimes \phi) \wedge (\phi_2 \rtimes \phi) \qquad\qquad \forall x : T.\phi_1 \rtimes \phi_2$$

$$(A \Rightarrow \phi_1) \rtimes \phi_2 = A \Rightarrow (\phi_1 \rtimes \phi_2)$$

The following Lemmas are used in the proofs of soundness and completeness to handle interleavings.

Lemma 16 (Shuffling correctness).

Let P_1, P_2, ϕ_1, ϕ_2*, if* $P_1 \models \phi_1$ *and* $P_2 \models \phi_2$ *and if* $\texttt{free}(\phi_1) \cap \texttt{free}(P_2) = \texttt{free}(\phi_2) \cap \texttt{free}(P_1) = \texttt{free}(P_1) \cap \texttt{free}(P_2) = \texttt{free}(\phi_1) \cap \texttt{free}(\phi_2) = \varnothing$*, then* $P_1 \mid P_2 \models \phi_1 \rtimes \phi_2$.

Conversely, if $P_1 \mid P_2 \models \phi_1 \rtimes \phi_2$, $\texttt{free}(\phi_1) \cap \texttt{free}(P_2) = \texttt{free}(\phi_2) \cap \texttt{free}(P_1) = \texttt{free}(P_1) \cap \texttt{free}(P_2) = \texttt{free}(\phi_1) \cap \texttt{free}(\phi_2) = \varnothing$, $\texttt{free}(\phi_1) \subseteq \texttt{free}(P_1)$ *and* $\texttt{free}(\phi_2) \subseteq \texttt{free}(P_2)$.

Lemma 17 (Description of free names). *If* $\mathcal{C}, \Gamma \vdash P : \Delta$ *then* $\texttt{free}(P) \subseteq \texttt{free}(\Delta) \cup \texttt{free}(\Gamma)$

Lemma 18 (Nature of an interleaving).

Let $\Delta = \{s_k[\mathtt{p}_k] : \mathtt{q}_k \overset{!}{?} \{ {}^{l_i(x_i)}_{e_i \mapsto l_i\langle e_i'\rangle(x_i)} \{A_i\}\langle E_i\rangle.T_{k,i}\}_{i\in I}\}_k$ *and* $\Gamma = \{a_j : \mathtt{I}(\mathcal{G}_j)\}_j$ *be well-formed, then the formula* $\texttt{Inter}(\Delta, \Gamma)$ *is equivalent to a formula guarded by several* $\forall$ *operators guarding a conjunction of formulae, each one starting with a modality, and this modalities are in bijection with the pairs of* $(s_k[{}^{\mathtt{p}_k, \mathtt{q}_k}_{\mathtt{q}_k, \mathtt{p}_k}], l_{k,i})$ *and* $(a_j, \varnothing)$.

Proposition 3 (Soundness). *If* $\mathcal{C}; \Gamma \vdash P \rhd \Delta$*, then:* $P, \sigma \models (\mathcal{C} \Rightarrow \texttt{Inter}(\Delta, \Gamma))$.

Completeness The erasing operator $\mathbf{Er}(\mathcal{L})$, which translates an asserted type into its unasserted counterpart is straightforwardly defined: we remove every assertion A from the local types. Unasserted typing rules for the judgements $\vdash P \rhd \Delta$ are easily deduced from the proof rules. Our completeness result is:

Proposition 4 (Completeness).

If $\vdash P \rhd \mathbf{Er}(\Delta)$ *and* $P, \sigma \models (\mathcal{C} \Rightarrow \texttt{Inter}(\Delta, \Gamma))$ *then* $\mathcal{C}; \Gamma \vdash P \rhd \Delta$.

Embedding to Pure HML We are able to embed a stateful satisfaction relation $P, \sigma \models \phi$ into a satisfaction relation $P' \models \phi'$ of a standard π-calculus with first-order values, by encoding the σ into a π-process:

$$\begin{array}{l}\|x_1 \mapsto v_1, \ldots, x_n \mapsto v_n\|_{\mathtt{p}} = \quad \overline{a_1}(v_1) \mid \ldots \mid \overline{a_n}(v_n) \mid \\ \quad !x_1(e).a_1(y_1)\ldots a_n(y_n).(\overline{a_1}(\texttt{eval}(e[y_1 \ldots y_n/x_1 \ldots x_n])) \mid \overline{a_2}(y_2) \mid \ldots \mid \overline{a_n}(y_n)) \mid \ldots \mid \\ \quad !x_n(e).a_1(y_1)\ldots a_n(y_n).(\overline{a_1}(y_1) \mid \ldots \mid \overline{a_{n-1}}(y_{n-1}) \mid \overline{a_n}(\texttt{eval}(e[y_1 \ldots y_n/x_1 \ldots x_n])))\end{array}$$

For each variable x_i in the domain of the stateσ, we add an output prefix emitting its content on the channel a_i and we add a replicated module that waits for an update e at x_i, then capture the value of all variables of the current state, replace the variable x_i by evaluating e by `eval`, and then makes available the other ones. Soundness and completeness allow us to state that HML formulae for pairs state/process can be seen as pure HML formulas on the π-processes.

Embedding for state σ is given by the following:

$$\begin{array}{c}\|x_1 \mapsto v_1, \ldots, x_n \mapsto v_n\|_{\mathtt{p}} = \overline{a_1}(v_1) \mid \ldots \mid \overline{a_n}(v_n) \mid \\ !x_1(e).a_1(y_1)\ldots a_n(y_n).(\overline{a_1}(\texttt{eval}(e\{y_1 \ldots y_n/x_1 \ldots x_n\})) \mid \overline{a_2}(y_2) \mid \ldots \mid \overline{a_n}(y_n)) \mid \\ \ldots \\ !x_n(e).a_1(y_1)\ldots a_n(y_n).(\overline{a_1}(y_1) \mid \ldots \mid \overline{a_{n-1}}(y_{n-1}) \mid \overline{a_n}(\texttt{eval}(e\{y_1 \ldots y_n/x_1 \ldots x_n\})))\end{array}$$

For each variable x_i in the domain of the stateσ, we add an output prefix emitting its content on the channel a_i and we add a replicated module that waits for an update e at x_i, then capture the value of all variable of the current state, replace the variable x_i by evaluating e w.r.t. the values of the state, and then makes available the other variables.

The embedding for the formula is given by the following:

$$\|[E]\phi\|_{\mathtt{p}} = [\|E\|_{\mathtt{p}}]\|\phi\|_{\mathtt{p}} \qquad \|A\|_{\mathtt{p}} = [\overline{x_1}(v_1)]\dots[\overline{x_n}(v_n)]A\{v_1,\dots,v_n/$$

$$x_1,\dots,x_n\} \text{ where the state variables of } A \text{ are } x_1,\dots,x_n$$

The following theorem ensures that the encoding is sound and complete.

Proposition 5 (Preciseness).
If $P,\sigma \models \phi$, *then* $\|P\|_{\mathtt{p}} \mid \|\sigma\|_{\mathtt{p}} \models \|\phi\|_{\mathtt{p}}$.
If $\|P\|_{\mathtt{p}} \mid \|\sigma\|_{\mathtt{p}} \models \|\phi\|_{\mathtt{p}}$ *then* $P,\sigma \models \phi$

Embedding Recursion Recursion is absent from the previous embeddings, but can actually be encoded, at the cost of much technical details, we give here a brief sketch of how we proceed. For this purpose, we add to our HML syntax the recursion operators, $\mu X.\phi$ and X (similar to the one present in the μ-calculus [15]).

The main difficulty lies in the interaction between interleaving and recursion: loops coming from different sessions can be interleaved in many different way, and the difficult task is to compute the finite formula which is equivalent to this interleaving.

As a small example consider the following session environment (interactions are replaced by integer labels): $s_1[\mathtt{p}_1] : \mu X.1.2.X, s_2[\mathtt{p}_2] : \mu Y.3.4.Y$. The simplest HML formula describing all possible interleavings is:

$$\mu A.([1]\mu B.([2]A \wedge [3]\mu C.([4]B \wedge [2]([1]C \wedge [4].A))) \wedge$$
$$[3]\mu D.([4].A \wedge [1]\mu E.([2]D \wedge [4]([2]A \wedge [3]E))))$$

We use the following method to obtain a matching HML formula. We use a translation through finite automata. Here is a sketch of the method, which takes as arguments a set session environment Δ:

1. Encode every session judgement $s_i[\mathtt{p}_i] : T_i$ of Δ independently into a formula ϕ_i, conforming to previous embedding and the definitions $\|\mu X.T\|^{s[\mathtt{p}]} = \mu X\|T\|^{s[\mathtt{p}]}$.
2. Translate every formula ϕ_i into a finite automata $\mathcal{A}_i$, one state corresponds to a point between two modalities or a μX in the formula, one transition correponds to either $[\ell](A \wedge [E]\circ)$ (output) or $[\ell](A \Rightarrow \circ)$ (input). Every automata is *directed* with a source state corresponding to the head of the formula and leaf states corresponding to recursion variables (or end of protocols).
3. Compute the automata $\mathcal{A}$, the parallel composition of all the $\mathcal{A}_i$, which is still *directed*.
4. Expand the automata $\mathcal{A}$, in order to obtain an equivalent branch automata, that is, an automata such that there is a root (the starting state) and transitions form a tree (back transitions are allowed but only on the same branch). This could be done by recursively replacing sub-automata with several copies of this sub-automata.

5. Translate back the automata into a formula, every state with more than two incoming transition is encoded as a recursion operator.

One our example, step 1 gives the formulas $\mu X.[1][2].X$ and $\mu Y.[3][4].Y$. Step 2 gives for each formula an automata with 2 states (initial and between $[1]$ (resp. $[3]$) and $[2]$ (resp. $[4]$)). Step 3 gives an automata with 4 states: the inital one, one after $[1]$, one after $[3]$, one after both $[1]$ and $[3]$. This automata is diamond-shaped, and, as a result, not tree-shaped. Step 4 yields an automata with 7 states, which is then translated in the formula described above.

The preciseness proof relies on the fact that the operation described in 3. and 4. give equivalent automata, and that two formulas translated to two equivalent automata are equivalent for the HML satisfaction relation.