Formally Verified Liveness with Synchronous Multiparty Session Types in Rocq

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Abstract

Multiparty session types (MPST) offer a framework for the description of communication-based protocols involving multiple participants. In the *top-down* approach to MPST, the communication pattern of the session is described using a *global type*. Then the global type is *projected* on to a *local type* for each participant, and the individual processes making up the session are type-checked against these projections. Typed sessions possess certain desirable properties such as *safety*, *deadlock-freedom* and *liveness* (also called *lock-freedom*).

In this work, we present the first mechanised proof of liveness for synchronous multiparty session types in the Rocq Proof Assistant. Building on recent work, we represent global and local types as coinductive trees using the paco library. We use a coinductively defined *subtyping* relation on local types together with another coinductively defined *plain-merge* projection relation relating local and global types . We then *associate* collections of local types, or *local type contexts*, with global types using this projection and subtyping relations, and prove an *operational correspondence* between a local type context and its associated global type. We then utilize this association relation to prove the safety and liveness of associated local type contexts and, consequently, the multiparty sessions typed by these contexts.

Besides clarifying the often informal proofs of liveness found in the MPST literature, our Rocq mechanisation also enables the certification of lock-freedom properties of communication protocols. Our contribution amounts to around 12K lines of Rocq code.

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1 Introduction

Multiparty session types [20] provide a type discipline for the correct-by-construction specification of message-passing protocols. Desirable protocol properties guaranteed by session types include *safety* (the labels and types of senders' payloads cohere with the capabilities of the receivers), *deadlock-freedom* (also called *progress* or *non-stuck property* [15]) (it is possible for the session to progress so long as it has at least one active participant), and *liveness* (also called *lock-freedom* [43] or *starvation-freedom* [9]) (if a process is waiting to send and receive then a communication involving it eventually happens).

There exists two common methodologies for multiparty session types. In the *bottom-up* approach, the individual processes making up the session are typed using a collection of *participants* and *local types*, that is, a *local type context*, and the properties of the session is examined by model-checking this local type context. Contrastingly, in the *top-down* approach sessions are typed by a *global type* that is related to the processes using endpoint *projections* and *subtyping*. The structure of the global type ensures that the desired properties are satisfied by the session. These two approaches have their advantages and disadvantages:

$$\begin{tikzcd}
\mathsf{G} \arrow[rr, dotted, "\rightarrow"] \arrow[d, dotted, "\upharpoonright"] \arrow[dr, "\sqsubseteq"] & & \mathsf{G}' \arrow[d, "\sqsubseteq"] \\
\mathsf{T} \arrow[r] \arrow[d, dotted, "\vdash_{\mathsf{P}}"] & \{(p_i, T_i) \mid i \in I\} = \Gamma \arrow[r, "\rightarrow"] \arrow[d, "\vdash_{\mathcal{M}}"] & \Gamma' \arrow[d, "\vdash_{\mathcal{M}}"] \\
\mathsf{P} \arrow[r] & \prod_i \mathsf{p}_i \triangleleft \mathsf{P}_i = \mathcal{M} \arrow[r, "\rightarrow"] & \mathcal{M}'
\end{tikzcd}$$

Figure 1 Design overview. The dotted lines correspond to relations inherited from [15] while the solid lines denote relations that are new, or substantially rewritten, in this paper.

the bottom-up approach is generally able to type more sessions, while type-checking and type-inferring in the top-down approach tend to be more efficient than model-checking the bottom-up system [42].

In this work, we present the Rocq [5] formalisation of a synchronous MPST that that ensures the aforementioned properties for typed sessions. Our type system uses an *association* relation ($\sqsubseteq$) [46, 34] defined using (coinductive plain) projection [40] and subtyping, in order to relate local type contexts and global types. This association relation ensures *operational correspondence* between the labelled transition system (LTS) semantics we define for local type contexts and global types. We then type ($\vdash_{\mathcal{M}}$) sessions using local type contexts that are associated with global types, which ensure that the local type context, and hence the session, is well-behaved in some sense. Whenever an associated local type context Γ types a session $\mathcal{M}$, our type system guarantees the following properties:

1. **Subject Reduction** (Theorem 6.2): If $\mathcal{M}$ can progress into $\mathcal{M}'$, then Γ can progress into Γ' such that Γ' types $\mathcal{M}'$.
2. **Session Fidelity** (Theorem 6.5): If Γ can progress into Γ', then $\mathcal{M}$ can progress into $\mathcal{M}'$ such that $\mathcal{M}'$ is typable by Γ'.
3. **Safety** (Theorem 6.7): If $\mathcal{M}$ can progress into $\mathcal{M}'$ by one or more communications, participant p in $\mathcal{M}'$ sends to participant q and q receives from p, then the labels of p and q cohere.
4. **Deadlock-Freedom** (Theorem 6.3): Either every participant in $\mathcal{M}$ has terminated, or $\mathcal{M}$ can progress.
5. **Liveness** (Theorem 6.16): If participant p attemps to communicate with participant q in $\mathcal{M}$, then $\mathcal{M}$ can progress (in possibly multiple steps) into a session $\mathcal{M}'$ where that communication has occured.

To our knowledge, this work presents the first mechanisation of liveness for multiparty session types in a proof assistant.

Our Rocq implementation builds upon the recent formalisation of subject reduction for MPST by Ekici et. al. [15], which itself is based on [18]. The methodology in [15] takes an equirecursive approach where an inductive syntactic global or local type is identified with the coinductive tree obtained by fully unfolding the recursion. It then defines a coinductive projection relation between global and local type trees, the LTS semantics for global type trees, and typing rules for the session calculus outlined in [18]. We extensively use these definitions and the lemmas concerning them, but we still depart from and extend [15] in numerous ways by introducing local typing contexts, their correspondence with global types and a new typing relation. Our addition to the code amounts to around 12K lines of Rocq code.

As with [15], our implementation heavily uses the parameterized coinduction technique of the paco [21] library. Namely, our liveness property is defined using possibly infinite *execution traces* which we represent as coinductive streams. The relevant predicates on these traces, such as fairness, are then defined using linear temporal logic (LTL)[35]. The LTL modalities eventually ($\diamond$) and always ($\Box$) can be expressed as least and greatest fixpoints respectively using expansion laws. This allows us to represent the properties that use these modalities as inductive and coinductive predicates in Rocq. This approach, together with the proof techniques provided by paco, results in compositional and clear proofs.

Outline. In Section 2 we define our session calculus and its LTS semantics. In Section 3 we introduce local and global type trees. In Section 4 we give LTS semantics to local type contexts and global types, and detail the association relation between them. In Section 5 we define safety and liveness for local type contexts, and prove that they hold for contexts associated with a global type tree. In Section 6 we give the typing rules for our session calculus, and prove the desired properties of these typable sessions.

2 The Session Calculus

We introduce the simple synchronous session calculus that our type system will be used on.

2.1 Processes and Sessions

▶ **Definition 2.1** (Expressions and Processes). *We define processes as follows:*

$$\mathsf{P} ::= \mathsf{p}!\ell(\mathsf{e}).\mathsf{P} \;\;|\;\; \sum_{i\in I} \mathsf{p}?\ell_i(x_i).\mathsf{P}_i \;\;|\;\; \text{if } \mathsf{e} \text{ then } \mathsf{P} \text{ else } \mathsf{P} \;\;|\;\; \mu \boldsymbol{X}.\mathsf{P} \;\;|\;\; \boldsymbol{X} \;\;|\;\; \boldsymbol{0}$$

where e *is an expression that can be a variable, a value such as* `true`, 0 *or* -3, *or a term built from expressions by applying the operators* `succ`, `neg`, $\neg$, *non-deterministic choice* $\oplus$ *and* $>$.

$\mathsf{p}!\ell(\mathsf{e}).\mathsf{P}$ is a process that sends the value of expression e with label ℓ to participant p, and continues with process P. $\sum_{i\in I} \mathsf{p}?\ell_i(x_i).\mathsf{P}_i$ is a process that may receive a value from p with any label ℓ_i where $i \in I$, binding the result to x_i and continuing with P_i, depending on which ℓ_i the value was received from. $\mathbf{X}$ is a recursion variable, $\mu\mathbf{X}.\mathsf{P}$ is a recursive process, if e then P else P is a conditional and $\mathbf{0}$ is a terminated process.

Processes can be composed in parallel into sessions.

▶ **Definition 2.2** (Multiparty Sessions). *Multiparty sessions are defined as follows.*

$$\mathcal{M} \;::=\; \mathsf{p} \triangleleft \mathsf{P} \quad | \quad (\mathcal{M} \mid \mathcal{M}) \quad | \quad \mathcal{O}$$

$\mathsf{p} \triangleleft \mathsf{P}$ denotes that participant p is running the process P, $|$ indicates parallel compositon. We write $\prod_{i\in I} \mathsf{p}_i \triangleleft \mathsf{P}_i$ to denote the session formed by p_i running P_i in parallel for all $i \in I$. $\mathcal{O}$ is an empty session with no participants, that is, the unit of parallel composition. In Rocq processes and sessions are defined with the inductive types `process` and `session`.

```
Inductive process : Type ≜
  | p_send : part → label → expr → process →
      process
  | p_recv : part → list(option process) → process
  | p_ite : expr → process → process → process
  | p_rec : process → process
  | p_var : nat → process
  | p_inact : process.
```

```
Inductive session: Type ≜
  | s_ind : part   → process → session
  | s_par : session → session → session
  | s_zero : session.
Notation "p '←−' P"  ≜  (s_ind p P) (at level 50, no
      associativity).
Notation "s1 '|||' s2" ≜  (s_par s1 s2) (at level 50, no
      associativity).
```

2.2 Structural Congruence and Operational Semantics

We define a structural congruence relation $\equiv$ on sessions which expresses the commutativity, associativity and unit of the parallel composition operator.

$$\textsc{[sc-sym]}\quad \mathsf{p} \triangleleft \mathsf{P} \mid \mathsf{q} \triangleleft \mathsf{Q} \equiv \mathsf{q} \triangleleft \mathsf{Q} \mid \mathsf{p} \triangleleft \mathsf{P} \qquad \textsc{[sc-assoc]}\quad (\mathsf{p} \triangleleft \mathsf{P} \mid \mathsf{q} \triangleleft \mathsf{Q}) \mid \mathsf{r} \triangleleft \mathsf{R} \equiv \mathsf{p} \triangleleft \mathsf{P} \mid (\mathsf{q} \triangleleft \mathsf{Q} \mid \mathsf{r} \triangleleft \mathsf{R})$$

$$\textsc{[sc-o]}\quad \mathsf{p} \triangleleft \mathsf{P} \mid \mathcal{O} \equiv \mathsf{p} \triangleleft \mathsf{P}$$

Table 1 Structural Congruence over Sessions

We now give the operational semantics for sessions by the means of a labelled transition system. We use labelled *reactive* semantics [43, 7] which doesn't contain explicit silent τ actions for internal reductions (that is, evaluation of if expressions and unfolding of recursion) while still considering β reductions up to those internal reductions by using an unfolding relation. This stands in contrast to the more standard semantics used in [15, 18, 43]. For the advantages of our approach see Remark 6.4.

In reactive semantics silent transitions are captured by an *unfolding* relation ($\Rrightarrow$), and β reductions are defined up to this unfolding (Table 2).

$$\textsc{[Unf-struct]}\ \frac{\mathcal{M} \equiv \mathcal{N}}{\mathcal{M} \Rrightarrow \mathcal{N}} \qquad \textsc{[Unf-rec]}\ \mathsf{p} \triangleleft \mu\mathbf{X}.\mathsf{P} \mid \mathcal{N} \Rrightarrow \mathsf{p} \triangleleft \mathsf{P}[\mu\mathbf{X}.\mathsf{P}/\mathbf{X}] \mid \mathcal{N} \qquad \textsc{[Unf-condT]}\ \frac{e \downarrow \text{true}}{\mathsf{p} \triangleleft \text{ if } e \text{ then } \mathsf{P} \text{ else } \mathsf{Q} \mid \mathcal{N} \Rrightarrow \mathsf{p} \triangleleft \mathsf{P} \mid \mathcal{N}}$$

$$\textsc{[Unf-condF]}\ \frac{e \downarrow \text{false}}{\mathsf{p} \triangleleft \text{ if } e \text{ then } \mathsf{P} \text{ else } \mathsf{Q} \mid \mathcal{N} \Rrightarrow \mathsf{p} \triangleleft \mathsf{Q} \mid \mathcal{N}} \qquad \textsc{[Unf-trans]}\ \frac{\mathcal{M} \Rrightarrow \mathcal{M}' \quad \mathcal{M}' \Rrightarrow \mathcal{N}}{\mathcal{M} \Rrightarrow \mathcal{N}}$$

Table 2 Unfolding of Sessions

$\mathcal{M} \Rrightarrow \mathcal{N}$ means that $\mathcal{M}$ can transition to $\mathcal{N}$ through some internal actions, that is, a reduction that doesn't involve a communication. We say that $\mathcal{M}$ *unfolds* to $\mathcal{N}$. In Rocq it's captured by the predicate `unfoldP : session → session → Prop`.

$$\textsc{[R-comm]}\ \frac{j \in I \quad e \downarrow v}{\mathsf{p} \triangleleft \sum_{i \in I} \mathsf{q}?\ell_i(x_i).\mathsf{P}_i \mid \mathsf{q} \triangleleft \mathsf{p}!\ell_j(\mathsf{e}).\mathsf{Q} \mid \mathcal{N} \xrightarrow{(\mathsf{p},\mathsf{q})\ell_j} \mathsf{p} \triangleleft \mathsf{P}_j[v/x_j] \mid \mathsf{q} \triangleleft \mathsf{Q} \mid \mathcal{N}}$$

$$\textsc{[R-unfold]}\ \frac{\mathcal{M} \Rrightarrow \mathcal{M}' \quad \mathcal{M}' \xrightarrow{\lambda} \mathcal{N}' \quad \mathcal{N}' \Rrightarrow \mathcal{N}}{\mathcal{M} \xrightarrow{\lambda} \mathcal{N}}$$

Table 3 Reactive Semantics of Sessions

Table 3 illustrates the rules for communicating transitons. [R-comm] captures communications between processes, and [R-unfold] lets us consider reductions up to unfoldings.

In Rocq, `betaP_lbl M lambda M'` denotes $\mathcal{M} \xrightarrow{\lambda} \mathcal{M}'$. We write $\mathcal{M} \to \mathcal{M}'$ if $\mathcal{M} \xrightarrow{\lambda} \mathcal{M}'$ for some λ, which is written `betaP M M'` in Rocq. We write $\to^*$ to denote the reflexive transitive closure of $\to$, which is called `betaRtc` in Rocq.

3 The Type System

We briefly recap the core definitions of local and global type trees, subtyping and projection from [18].

3.1 Local Types and Type Trees

We start by defining the sorts that will be used to type expressions, and local types that will be used to type single processes.

▶ **Definition 3.1** (Sorts). *Sorts are defined as follows:*

$$S ::= \ \texttt{int} \ | \ \texttt{bool} \ | \ \texttt{nat}$$

```
Inductive sort: Type ≜
  | sbool: sort
  | sint : sort
  | snat : sort.
```

▶ **Definition 3.2.** *Local types are defined inductively with the following syntax:*

$$\mathbb{T} ::= \ \texttt{end} \ | \ \mathsf{p}\oplus\{\ell_i(S_i).\mathbb{T}_i\}_{i\in I} \ | \ \mathsf{p}\&\{\ell_i(S_i).\mathbb{T}_i\}_{i\in I} \ | \ \boldsymbol{t} \ | \ \mu\boldsymbol{t}.\mathbb{T}$$

Informally, in the above definition, `end` represents a role that has finished communicating. $\mathsf{p}\oplus\{\ell_i(S_i).\mathbb{T}_i\}_{i\in I}$ denotes a role that may, from any $i \in I$, receive a value of sort S_i with message label ℓ_i and continue with $\mathbb{T}_i$. Similarly, $\mathsf{p}\&\{\ell_i(S_i).\mathbb{T}_i\}_{i\in I}$ represents a role that may choose to send a value of sort S_i with message label ℓ_i and continue with $\mathbb{T}_i$ for any $i \in I$. $\mu\mathbf{t}.\mathbb{T}$ represents a recursive type where $\mathbf{t}$ is a type variable. We assume that the indexing sets I are always non-empty. We also assume that recursion is always guarded.

We employ an equirecursive approach based on the standard techniques from [33] where $\mu\mathbf{t}.\mathbb{T}$ is considered to be equivalent to its unfolding $\mathbb{T}[\mu\mathbf{t}.\mathbb{T}/\mathbf{t}]$. This enables us to identify a recursive type with the possibly infinite local type tree obtained by fully unfolding its recursive subterms.

▶ **Definition 3.3.** *Local type trees are defined coinductively with the following syntax:*

$$\begin{aligned} \mathsf{T} ::= \ & \texttt{end} \\ | \ & \mathsf{p}\&\{\ell_i(S_i).\mathsf{T}_i\}_{i\in I} \\ | \ & \mathsf{p}\oplus\{\ell_i(S_i).\mathsf{T}_i\}_{i\in I} \end{aligned}$$

```
CoInductive ltt: Type ≜
  | ltt_end : ltt
  | ltt_recv: part → list (option(sort*ltt)) → ltt
  | ltt_send: part → list (option(sort*ltt)) → ltt.
```

In Rocq we represent the continuations using a `list` of `option` types. In a continuation `gcs : list (option(sort*ltt))`, index `k` (using zero-indexing) being equal to `Some (s_k, T_k)` means that $\ell_k(S_k).\mathsf{T}_k$ is available in the continuation. Similarly index `k` being equal to `None` or being out of bounds of the list means that the message label ℓ_k is not present in the continuation.

▶ Remark 3.4. Note that Rocq allows us to create types such as `ltt_send q []` which don't correspond to well-formed local types as the continuation is empty. In our implementation we define a predicate `wfltt : ltt → Prop` capturing that all the continuations in the local type tree are non-empty. Henceforth we assume that all local types we mention satisfy this property.

We omit the details of the translation between local types and local type trees, the technicalities of our approach is explained in [18], and the Rocq implementation of translation is detailed in [15]. From now on we work exclusively on local type trees. Also, as done in [15], we assume coinductive extensionality and consider isomorphic type trees to be equal.

3.2 Subtyping

We define the subsorting relation on sorts and the subtyping relation on local type trees.

▶ **Definition 3.5** (Subsorting and Subtyping). *Subsorting $\leq$ is the least reflexive binary relation that satisfies* `nat` $\leq$ `int`. *Subtyping $\leqslant$ is the largest relation between local type trees coinductively defined by the following rules:*

$$\frac{}{\texttt{end} \leqslant \texttt{end}}\ [\textsc{sub-end}] \qquad \frac{\forall i \in I: \quad S'_i \leq S_i \quad \mathsf{T}_i \leqslant \mathsf{T}'_i}{\mathsf{p}\&\{\ell_i(S_i).\mathsf{T}_i\}_{i\in I\cup J} \leqslant \mathsf{p}\&\{\ell_i(S'_i).\mathsf{T}'_i\}_{i\in I}}\ [\textsc{sub-in}]$$

$$\frac{\forall i \in I: \quad S_i \leq S'_i \quad \mathsf{T}_i \leqslant \mathsf{T}'_i}{\mathsf{p}\oplus\{\ell_i(S_i).\mathsf{T}_i\}_{i\in I} \leqslant \mathsf{p}\oplus\{\ell_i(S'_i).\mathsf{T}'_i\}_{i\in I\cup J}}\ [\textsc{sub-out}]$$

Intuitively, $\mathsf{T}_1 \leqslant \mathsf{T}_2$ means that a role of type T_1 can be supplied anywhere a role of type T_2 is needed. [SUB-IN] captures the fact that we can supply a role that is able to receive more labels than specified, and [SUB-OUT] captures that we can supply a role that has fewer labels available to send. Note the contraviance of the sorts in [SUB-IN], if the supertype demands the ability to receive an `nat` then the subtype can receive `nat` or `int`.

In Rocq, the subtyping relation `subtypeC : ltt → ltt → Prop` is expressed as a greatest fixpoint using the `Paco` library [21], for details of we refer to [18].

3.3 Global Types and Type Trees

While local types specify the behaviour of one role in a protocol, global types give a bird's eye view of the whole protocol.

▶ **Definition 3.6** (Global type). *We define global types inductively as follows:*

$$\mathbb{G} ::= \quad \texttt{end} \ \mid\ \mathsf{p} \to \mathsf{q} : \{\ell_i(S_i).\mathbb{G}_i\}_{i\in I} \ \mid\ \boldsymbol{t} \ \mid\ \mu\boldsymbol{t}.\mathbb{G}$$

We further inductively define the function $\texttt{pt}(\mathbb{G})$ *that denotes the participants of type* $\mathbb{G}$*:*

$$\texttt{pt}(\texttt{end}) = \texttt{pt}(\boldsymbol{t}) = \emptyset$$

$$\texttt{pt}(\mathsf{p} \to \mathsf{q} : \{\ell_i(S_i).\mathbb{G}_i\}_{i\in I}) = \{\mathsf{p}, \mathsf{q}\} \cup \bigcup_{i\in I} \texttt{pt}(\mathbb{G}_i)$$

$$\texttt{pt}(\mu\boldsymbol{T}.\mathbb{G}) = \texttt{pt}(\mathbb{G})$$

`end` denotes a protocol that has ended, $\mathsf{p} \to \mathsf{q} : \{\ell_i(S_i).\mathbb{G}_i\}_{i\in I}$ denotes a protocol where for any $i \in I$, participant p may send a value of sort S_i to another participant q via message label ℓ_i, after which the protocol continues as $\mathbb{G}_i$.

As in the case of local types, we adopt an equirecursive approach and work exclusively on possibly infinite global type trees.

▶ **Definition 3.7** (Global type trees). *We define global type trees coinductively as follows:*

$$\mathsf{G} ::= \quad \mathsf{end} \ \mid \ \mathsf{p} \to \mathsf{q} : \{\ell_i(S_i).\mathsf{G}_i\}_{i \in I}$$

```
CoInductive gtt: Type ≜
| gtt_end    : gtt

| gtt_send   : part → part → list (option (sort*gtt)) →
    gtt.
```

We extend the function $\mathtt{pt}$ onto trees by defining $\mathtt{pt}(\mathsf{G}) = \mathtt{pt}(\mathbb{G})$ where the global type $\mathbb{G}$ corresponds to the global type tree G. Technical details of this definition such as well-definedness can be found in [15, 18].

In Rocq $\mathtt{pt}$ is captured with the predicate `isgPartsC : part → gtt → Prop`, where `isgPartsC p G` denotes $\mathsf{p} \in \mathtt{pt}(\mathsf{G})$.

3.4 Projection

We now define coinductive projections with plain merging (see [42] for a survey of other notions of merge).

▶ **Definition 3.8** (Projection). *The projection of a global type tree onto a participant* r *is the largest relation* $\upharpoonright_{\mathsf{r}}$ *between global type trees and local type trees such that, whenever* $\mathsf{G} \upharpoonright_{\mathsf{r}} \mathsf{T}$*:*

- $\mathsf{r} \notin \mathtt{pt}\{\mathsf{G}\}$ *implies* $\mathsf{T} = \mathsf{end}$*;* [Proj-End]
- $\mathsf{G} = \mathsf{p} \to \mathsf{r} : \{\ell_i(S_i).\mathsf{G}_i\}_{i \in I}$ *implies* $\mathsf{T} = \mathsf{p}\&\{\ell_i(S_i).\mathsf{T}_i\}_{i \in I}$ *and* $\forall i \in I, \mathsf{G} \upharpoonright_{\mathsf{r}} \mathsf{T}_i$ [Proj-In]
- $\mathsf{G} = \mathsf{r} \to \mathsf{q} : \{\ell_i(S_i).\mathsf{G}_i\}_{i \in I}$ *implies* $\mathsf{T} = \mathsf{q}\oplus\{\ell_i(S_i).\mathsf{T}_i\}_{i \in I}$ *and* $\forall i \in I, \mathsf{G} \upharpoonright_{\mathsf{r}} \mathsf{T}_i$ [Proj-Out]
- $\mathsf{G} = \mathsf{p} \to \mathsf{q} : \{\ell_i(S_i).\mathsf{G}_i\}_{i \in I}$ *and* $\mathsf{r} \notin \{\mathsf{p}, \mathsf{q}\}$ *implies that there are* $\mathsf{T}_i, i \in I$ *such that* $\mathsf{T} = \bigsqcap_{i \in I} \mathsf{T}_i$ *and* $\forall i \in I, \mathsf{G} \upharpoonright_{\mathsf{r}} \mathsf{T}_i$ [Proj-Cont]

where $\bigsqcap$ *is the plain merging operator, defined as*

$$\mathsf{T}_1 \bigsqcap \mathsf{T}_2 = \begin{cases} \mathsf{T}_1 & \text{if } \mathsf{T}_1 = \mathsf{T}_2 \\ \text{undefined} & \text{otherwise} \end{cases}$$

Informally, the projection of a global type tree G onto a participant r extracts a specification for participant r from the protocol whose bird's-eye view is given by G. [Proj-End] expresses that if r is not a participant of G then r does nothing in the protocol. [Proj-In] and [Proj-Out] handle the cases where r is involved in a communication in the root of G. [Proj-Cont] says that, if r is not involved in the root communication of G, then the only way it knows its role in the protocol is if there is a role for it that works no matter what choices p and q make in their communication. This "works no matter the choices of the other participants" property is captured by the merge operations.

In Rocq, projection is defined as a `Paco` greatest fixpoint as the relation `projectionC : gtt → part → ltt → Prop`.

We further have the following fact about projections that lets us regard it as a partial function:

▶ **Lemma 3.9.** *If* `projectionC G p T` *and* `projectionC G p T'` *then* `T = T'`.

We write $\mathsf{G} \upharpoonright \mathsf{r} = \mathsf{T}$ when $\mathsf{G} \upharpoonright_{\mathsf{r}} \mathsf{T}$. Furthermore we will be frequently be making assertions about subtypes of projections of a global type e.g. $\mathsf{T} \leqslant \mathsf{G} \upharpoonright \mathsf{r}$. In our Rocq implementation we define the predicate `issubProj : ltt → gtt → part → Prop` as a shorthand for this.

3.5 Balancedness, Global Tree Contexts and Grafting

We introduce an important constraint on the types of global type trees we will consider, balancedness.

▶ **Definition 3.10** (Balanced Global Type Trees). *A global tree* G *is balanced if for any subtree* G' *of* G*, there exists* k *such that for all* $\mathsf{p} \in \mathtt{pt}(\mathsf{G}')$*,* p *occurs on every path from the root of* G' *of length at least* k*.*

We omit the technical details of this definition and the Rocq implementation, they can be found in [18] and [15].

Intutively, balancedness is a regularity condition that imposes a notion of *liveness* on the protocol described by the global type tree. Indeed, our liveness results in Section 6 hold only for balanced global types. Another reason for formulating balancedness is that it allows us to use the "grafting" technique, turning proofs by coinduction on infinite trees to proofs by induction on finite global type tree contexts.

▶ **Definition 3.11** (Global Type Tree Context). *Global type tree contexts are defined inductively with the following syntax:*

$$\mathcal{G} ::= \quad \mathsf{p} \to \mathsf{q} : \{\ell_i(S_i).\mathcal{G}_i\}_{i \in I} \;\;\big|\;\; [\;]_i$$

```
Inductive gtth: Type ≜
  | gtth_hol    : fin → gtth
  | gtth_send   : part → part → list (option (sort *
       gtth)) → gtth.
```

We additionally define `pt` *and* `ishParts` *on contexts analogously to* `pt` *and* `isgPartsC` *on trees.*

A global type tree context can be thought of as the finite prefix of a global type tree, where holes $[\;]_i$ indicate the cutoff points. Global type tree contexts are related to global type trees with the grafting operation.

▶ **Definition 3.12** (Grafting). *Given a global type tree context* $\mathcal{G}$ *whose holes are in the indexing set* I *and a set of global types* $\{\mathsf{G}_i\}_{i \in I}$*, the grafting* $\mathcal{G}[\mathsf{G}_i]_{i \in I}$ *denotes the global type tree obtained by substituting* $[\;]_i$ *with* G_i *in* $\mathcal{G}$*.*

In Rocq the indexed set $\{\mathsf{G}_i\}_{i \in I}$ *is represented using a* `list (option gtt)`*. Grafting is expressed with the inductive relation* `typ_gtth : list (option gtt) → gtth → gtt → Prop`*.* `typ_gtth gs gcx gt` *means that the grafting of the set of global type trees* `gs` *onto the context* `gcx` *results in the tree* `gt`*.*

Furthermore, we have the following lemma that relates global type tree contexts to balanced global type trees.

▶ **Lemma 3.13** (Proper Grafting Lemma, [15]). *If* `G` *is a balanced global type tree and* `isgPartsC p G`*, then there is a global type tree context* `Gctx` *and an option list of global type trees* `gs` *such that* `typ_gtth gs Gctx G`*,* `~ ishParts p Gctx` *and every* `Some` *element of* `gs` *is of shape* `gtt_end`*,* `gtt_send p q` *or* `gtt_send q p`*.*

3.13 enables us to represent a coinductive global type tree featuring participant `p` as the grafting of a context that doesn't contain `p` with a list of trees that are all of a certain structure. If `typ_gtth gs Gctx G`, `~ ishParts p Gctx` and every `Some` element of `gs` is of shape `gtt_end`, `gtt_send p q` or `gtt_send q p`, then we call the pair `gs` and `Gctx` as the `p`-grafting of `G`, expressed in Rocq as `typ_p_gtth gs Gctx p G`. When we don't care about the contents of `gs` we may just say that `G` is `p`-grafted by `Gctx`.

▶ Remark 3.14. From now on, all the global type trees we will be referring to are assumed to be balanced. When talking about the Rocq implementation, any `G : gtt` we mention is assumed to satisfy the predicate `wfgC G`, expressing that `G` corresponds to some global type and that `G` is balanced. Furthermore, we will often require that a global type is projectable onto all its participants. This is captured by the predicate `projectableA G = ∀ p, ∃ T, projectionC G p T`. As with `wfgC`, we will be assuming that all types we mention are projectable.

4 Semantics of Types

In this section we introduce local type contexts, and define Labelled Transition System semantics on these constructs.

4.1 Typing Contexts

We start by defining typing contexts as finite mappings of participants to local type trees.

▶ **Definition 4.1** (Typing Contexts).

$$\Gamma \ ::= \ \emptyset \ \mid \ \Gamma, \mathsf{p} : \mathsf{T}$$

```
Module M ≜ MMaps.RBT.Make(Nat).
Module MF ≜ MMaps.Facts.Properties Nat M.
Definition tctx: Type ≜ M.t ltt.
```

Intuitively, p : T means that participant p is associated with a process that has the type tree T. We write $\mathrm{dom}(\Gamma)$ to denote the set of participants occuring in Γ. We write $\Gamma(\mathsf{p})$ for the type of p in Γ. We define the composition Γ_1, Γ_2 iff $\mathrm{dom}(\Gamma_1) \cap \mathrm{dom}(\Gamma_2) = \emptyset$.

In the Rocq implementation we implement local typing contexts as finite maps of participants, which are represented as natural numbers, and local type trees. We use the red-black tree based finite map implementation of the MMaps library [28].

▶ Remark 4.2. From now on, we assume the all the types in the local type contexts always have non-empty continuations. In Rocq terms, if `T` is in context `gamma` then `wfltt T` holds. This is expressed by the predicate `wfltt: tctx → Prop`.

4.2 Local Type Context Reductions

We now give LTS semantics to local typing contexts, for which we first define the transition labels.

▶ **Definition 4.3** (Transition labels). *A transition label α has the following form:*

$$\begin{array}{lll} \alpha ::= \mathsf{p} : \mathsf{q} \& \ell(S) & & (\mathsf{p} \textit{ receives a value of sort } S \textit{ from } \mathsf{q} \textit{ with message label } \ell) \\ \quad \mid \ \mathsf{p} : \mathsf{q} \oplus \ell(S) & & (\mathsf{p} \textit{ sends a value of sort } S \textit{ to } \mathsf{q} \textit{ with message label } \ell\ ) \\ \quad \mid \ (\mathsf{p}, \mathsf{q})\ell & & (\textit{A synchronized communication from } \mathsf{p} \textit{ to } \mathsf{q} \textit{ occurs via message label } \ell) \end{array}$$

In Rocq they are defined as follows:

```
Notation opt_lbl ≜ nat.
Inductive label: Type ≜
  | lrecv: part → part → option sort → opt_lbl → label
  | lsend: part → part → option sort → opt_lbl → label
  | lcomm: part → part → opt_lbl → label.
```

Next we define labelled transitions for local type contexts.

▶ **Definition 4.4** (Typing context reductions). *The typing context transition $\xrightarrow{\alpha}$ is defined inductively by the following rules:*

$$\frac{k \in I}{\mathsf{p} : \mathsf{q}\&\{\ell_i(S_i).\mathsf{T}_i\}_{i\in I} \xrightarrow{\mathsf{p}:\mathsf{q}\&\ell_k(S_k)} \mathsf{p} : \mathsf{T}_k}\ [\,\Gamma\text{-}\&]$$

$$\frac{k \in I}{\mathsf{p} : \mathsf{q}\oplus\{\ell_i(S_i).\mathsf{T}_i\}_{i\in I} \xrightarrow{\mathsf{p}:\mathsf{q}\oplus\ell_k(S_k)} \mathsf{p} : \mathsf{T}_k}\ [\,\Gamma\text{-}\oplus] \qquad \frac{\Gamma \xrightarrow{\alpha} \Gamma'}{\Gamma, \mathsf{p} : \mathsf{T} \xrightarrow{\alpha} \Gamma', \mathsf{p} : \mathsf{T}}\ [\Gamma\text{-},]$$

$$\frac{\Gamma_1 \xrightarrow{\mathsf{p}:\mathsf{q}\oplus\ell(S)} \Gamma_1' \qquad \Gamma_2 \xrightarrow{\mathsf{q}:\mathsf{p}\&\ell(S')} \Gamma_2' \qquad S \leq S'}{\Gamma_1, \Gamma_2 \xrightarrow{(\mathsf{p},\mathsf{q})\ell} \Gamma_1', \Gamma_2'}\ [\Gamma\text{-}\oplus\&]$$

We write $\Gamma \xrightarrow{\alpha}$ if there exists Γ' such that $\Gamma \xrightarrow{\alpha} \Gamma'$. We define a reduction $\Gamma \to \Gamma'$ that holds iff $\Gamma \xrightarrow{(\mathsf{p},\mathsf{q})\ell} \Gamma'$ for some p, q, ℓ. We write $\Gamma \to$ iff $\Gamma \to \Gamma'$ for some Γ'. We write $\to^*$ for the reflexive transitive closure of $\to$.

$[\,\Gamma\text{-}\oplus]$ and $[\,\Gamma\text{-}\&]$, express a single participant sending or receiving. $[\Gamma\text{-}\oplus\&]$ expresses a synchronized communication where one participant sends while another receives, and they both progress with their continuation. $[\Gamma\text{-},]$ shows how to extend a context.

In Rocq typing context reductions are defined the following way:

```
Inductive tctxR: tctx → label → tctx → Prop ≜
  | Rsend: ∀ p q xs n s T,
           p ≠ q →
           onth n xs = Some (s, T) →
           tctxR (M.add p (ltt_send q xs) M.empty) (lsend p q (Some s) n) (M.add p T M.empty)
  | Rrecv: ...
  | Rcomm: ∀ p q g1 g1' g2 g2' s s' n (H1: MF.Disjoint g1 g2) (H2: MF.Disjoint g1' g2'),
           p ≠ q →
           tctxR g1 (lsend p q (Some s) n) g1'  →
           tctxR g2 (lrecv q p (Some s') n) g2' →
           subsort s s' →
           tctxR (disj_merge g1 g2 H1) (lcomm p q n) (disj_merge g1' g2' H2)
  | RvarI: ∀ g l g' p T,
           tctxR g l g' →
           M.mem p g = false →
           tctxR (M.add p T g) l (M.add p T g')
  | Rstruct: ∀ g1 g1' g2 g2' l, tctxR g1' l g2' →
    M.Equal g1 g1' →
    M.Equal g2 g2' →
    tctxR g1 l g2.
```

`Rsend`, `Rrecv` and `RvarI` are straightforward translations of $[\,\Gamma\text{ - }\&]$, $[\,\Gamma\text{ - }\oplus]$ and $[\Gamma\text{ -},]$. `Rcomm` captures $[\Gamma - \oplus\&]$ using the `disj_merge` function we defined for the compositions, and requires a proof that the contexts given are disjoint to be applied. `RStruct` captures the indistinguishability of local contexts under the `M.Equal` predicate from the MMaps library. We give an example to illustrate typing context reductions.

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▶ **Example 4.5.** Let

$$\mathsf{T}_\mathsf{p} = \mathsf{q}\oplus\{\ell_0(\mathtt{int}).\mathsf{T}_\mathsf{p}\ ,\ \ell_1(\mathtt{int}).\mathsf{end}\}$$
$$\mathsf{T}_\mathsf{q} = \mathsf{p}\&\{\ell_0(\mathtt{int}).\mathsf{T}_\mathsf{q}\ ,\ \ell_1(\mathtt{int}).\mathsf{r}\oplus\{\ell_2(\mathtt{int}).\mathsf{end}\}\}$$
$$\mathsf{T}_\mathsf{r} = \mathsf{q}\&\{\ell_2(\mathtt{int}).\mathsf{end}\}$$

and $\Gamma = \{\mathsf{p} : \mathsf{T}_\mathsf{p},\ \mathsf{q} : \mathsf{T}_\mathsf{q},\ \mathsf{r} : \mathsf{T}_\mathsf{r}\}$. We have the reductions $\Gamma \xrightarrow{\mathsf{p}:\mathsf{q}\oplus\ell_0(\mathtt{int})} \Gamma$ and $\Gamma \xrightarrow{\mathsf{q}:\mathsf{p}\&\ell_0(\mathtt{int})} \Gamma$, which synchronise to give the reduction and $\Gamma \xrightarrow{(\mathsf{p},\mathsf{q})\ell_0} \Gamma$. Similarly via synchronised communication of p and q via message label ℓ_1 we get $\Gamma \xrightarrow{(\mathsf{p},\mathsf{q})\ell_1} \Gamma'$ where Γ' is defined as $\{\mathsf{p} : \mathsf{end},\ \mathsf{q} : \mathsf{r}\oplus\{\ell_2(\mathtt{int}).\mathsf{end}\}, \mathsf{r} : \mathsf{T}_\mathsf{r}\}$.

In Rocq, Γ is defined the following way:

```
Definition prt_p≜0.
Definition prt_q≜1.
Definition prt_r≜2.
CoFixpoint T_p ≜ ltt_send prt_q [Some (sint,T_p); Some (sint,ltt_end); None].
CoFixpoint T_q ≜ ltt_recv prt_p [Some (sint,T_q); Some (sint, ltt_send prt_r [None;None;Some (sint,ltt_end)]); None].
Definition T_r ≜ ltt_recv prt_q [None;None; Some (sint,ltt_end)].
Definition gamma ≜ M.add prt_p T_p (M.add prt_q T_q (M.add prt_r T_r M.empty)).
```

Now $\Gamma \xrightarrow{(\mathsf{p},\mathsf{q})\ell_0} \Gamma$ can be expressed as `tctxR gamma (lsend prt_p prt_q (Some sint) 0) gamma`.

4.3 Global Type Reductions

As with local typing contexts, we can also define reductions for global types.

▶ **Definition 4.6** (Global type reductions). *The global type transition $\xrightarrow{\alpha}$ is defined coinductively as follows.*

$$\frac{k \in I}{\mathsf{p} \to \mathsf{q} : \{\ell_i(S_i).\mathsf{G}_i\}_{i\in I} \xrightarrow{(\mathsf{p},\mathsf{q})\ell_k} \mathsf{G}_k} \ [\text{GR-}\oplus\&]$$

$$\frac{\forall i \in I \ \ \mathsf{G}_i \xrightarrow{\alpha} \mathsf{G}'_i \qquad \text{subject}(\alpha) \cap \{\mathsf{p},\mathsf{q}\} = \emptyset \qquad \forall i \in I \ \ \{\mathsf{p},\mathsf{q}\} \subseteq \mathsf{pt}\{\mathsf{G}_i\}}{\mathsf{p} \to \mathsf{q} : \{\ell_i(S_i).\mathsf{G}_i\}_{i\in I} \xrightarrow{\alpha} \mathsf{p} \to \mathsf{q} : \{\ell_i(S_i).\mathsf{G}'_i\}_{i\in I}} \ [\text{GR-}\textsc{Ctx}]$$

[GR-⊕&] says that a global type tree with root $\mathsf{p} \to \mathsf{q}$ can transition to any of its children corresponding to the message label choosen by p. [GR-Ctx] says that if the subjects of α are disjoint from the root and all its children can transition via α, then the whole tree can also transition via α, with the root remaining the same and just the subtrees of its children transitioning.

In Rocq global type reductions are expressed using the coinductively defined predicate `gttstepC`. For example, $\mathsf{G} \xrightarrow{(\mathsf{p},\mathsf{q})\ell_k} \mathsf{G}'$ translates to `gttstepC G G' p q k`. We refer to [15] for details.

4.4 Association Between Local Type Contexts and Global Types

We have defined local type contexts which specifies protocols bottom-up by directly describing the roles of every participant, and global types, which give a top-down view of the whole protocol, and the transition relations on them. We now relate these local and global definitions by defining *association* between local type context and global types.

▶ **Definition 4.7** (Association). *A local typing context Γ is associated with a global type tree G, written $\Gamma \sqsubseteq \mathsf{G}$, if the following hold:*

- *For all* $\mathsf{p} \in \mathsf{pt}(\mathsf{G})$, $\mathsf{p} \in \mathrm{dom}(\Gamma)$ *and* $\Gamma(\mathsf{p}) \leqslant \mathsf{G} \upharpoonright \mathsf{p}$.
- *For all* $\mathsf{p} \notin \mathsf{pt}(\mathsf{G})$, *either* $\mathsf{p} \notin \mathrm{dom}(\Gamma)$ *or* $\Gamma(\mathsf{p}) = \mathsf{end}$.

In Rocq this is defined with the following:

```
Definition assoc (g: tctx) (gt:gtt) ≜
    ∀ p, (isgPartsC p gt  →  ∃ Tp, M.find p g=Some Tp ∧
      issubProj Tp gt p) ∧
       (~ isgPartsC p gt  →  ∀ Tpx, M.find p g = Some Tpx  →  Tpx=ltt_end).
```

Informally, $\Gamma \sqsubseteq \mathsf{G}$ says that the local type trees in Γ obey the specification described by the global type tree G.

▶ **Example 4.8.** In Example 4.5, we have that $\Gamma \sqsubseteq \mathsf{G}$ where

$$\mathsf{G} := \mathsf{p} \to \mathsf{q} : \{\ell_0(\mathtt{int}).\mathsf{G}, \ell_1(\mathtt{int}).\mathsf{q} \to \mathsf{r} : \{\ell_2(\mathtt{int}).\mathsf{end}\}\}$$

In fact, we have $\Gamma(\mathsf{s}) = \mathsf{G} \upharpoonright \mathsf{s}$ for $\mathsf{s} \in \{\mathsf{p}, \mathsf{q}, \mathsf{r}\}$. Similarly, we have $\Gamma' \sqsubseteq \mathsf{G}'$ where

$$\mathsf{G}' := \mathsf{q} \to \mathsf{r} : \{\ell_2(\mathtt{int}).\mathsf{end}\}$$

It is desirable to have the association be preserved under local type context and global type reductions, that is, when one of the associated constructs "takes a step" so should the other. We formalise this property with soundness and completeness theorems.

▶ **Theorem 4.9** (Soundness of Association). *If* `assoc gamma G` *and* `gttstepC G G' p q ell`*, then there is a local type context* `gamma'`*, a global type tree* `G''` *and a message label* `ell'` *such that* `gttStepC G G'' p q ell'`*,* `assoc gamma' G''` *and* `tctxR gamma (lcomm p q ell') gamma'`*.*

▶ **Theorem 4.10** (Completeness of Association). *If* `assoc gamma G` *and* `tctxR gamma (lcomm p q ell) gamma'`*, then there exists a global type tree* `G'` *such that* `assoc gamma' G'` *and* `gttstepC G G' p q ell`*.*

▶ Remark 4.11. Note that in the statement of soundness we allow the message label for the local type context reduction to be different to the message label for the global type reduction. This is because our use of subtyping in association causes the entries in the local type context to be less expressive than the types obtained by projecting the global type. For example consider

$$\Gamma = \mathsf{p} : \mathsf{q}\oplus\{\ell_0(\mathtt{int}).\mathsf{end}\},\ \mathsf{q} : \mathsf{p}\&\{\ell_0(\mathtt{int}).\mathsf{end}, \ell_1(\mathtt{int}).\mathsf{end}\}$$

and

$$\mathsf{G} = \mathsf{p} \to \mathsf{q} : \{\ell_0(\mathtt{int}).\mathsf{end}, \ell_1(\mathtt{int}).\mathsf{end}\}$$

We have $\Gamma \sqsubseteq \mathsf{G}$ and $\mathsf{G} \xrightarrow{(\mathsf{p},\mathsf{q})\ell_1}$. However $\Gamma \xrightarrow{(\mathsf{p},\mathsf{q})\ell_1}$ is not a valid transition. Note that soundness still requires that $\Gamma \xrightarrow{(\mathsf{p},\mathsf{q})\ell_x}$ for some x, which is satisfied in this case by the valid transition $\Gamma \xrightarrow{(\mathsf{p},\mathsf{q})\ell_0}$.

5 Properties of Local Type Contexts

We now use the LTS semantics to define some desirable properties on type contexts and their reduction sequences. Namely, we formulate safety, liveness and fairness properties based on the definitions in [46].

5.1 Safety

We start by defining safety:

▶ **Definition 5.1** (Safe Type Contexts). *We define* safe *coinductively as the largest set of type contexts such that whenever we have* $\Gamma \in \mathsf{safe}$*:*

$$\Gamma \xrightarrow{\mathsf{p}:\mathsf{q}\oplus\ell(S)} \textit{ and } \Gamma \xrightarrow{\mathsf{q}:\mathsf{p}\&\ell'(S')} \textit{ implies } \Gamma \xrightarrow{(\mathsf{p},\mathsf{q})\ell} \qquad \text{[S-\&}\oplus\text{]}$$

$$\Gamma \to \Gamma' \textit{ implies } \Gamma' \in \mathsf{safe} \qquad \text{[S-}\to\text{]}$$

We write $\mathsf{safe}(\Gamma)$ *if* $\Gamma \in \mathsf{safe}$*.*

Informally, safety says that if p and q communicate with each other and p requests to send a value using message label ℓ, then q should be able to receive that message label. Furthermore, this property should be preserved under any typing context reductions. Being a coinductive property, to show that $\mathsf{safe}(\Gamma)$ it suffices to give a set φ such that $\Gamma \in \varphi$ and φ satisfies [S-&⊕] and [S-→]. This amounts to showing that every element of Γ' of the set of reducts of Γ, defined $\varphi := \{\Gamma' \mid \Gamma \to^* \Gamma'\}$, satisfies [S-&⊕]. We illustrate this with some examples:

▶ **Example 5.2.** Let $\Gamma_A = \mathsf{p} : \mathsf{end}$, then Γ_A is safe: the set of reducts is $\{\Gamma_A\}$ and this set respects [S- ⊕&] as its elements can't reduce, and it respects [S-→] as it's closed with respect to $\to$.

Let $\Gamma_B = \mathsf{p} : \mathsf{q}{\oplus}\{\ell_0(\mathtt{int}).\mathsf{end}\}, \mathsf{q} : \mathsf{p}\&\{\ell_0(\mathtt{nat}).\mathsf{end}\}$. Γ_B is not safe as as we have $\Gamma_B \xrightarrow{\mathsf{p}:\mathsf{q}\oplus\ell_0}$ and $\Gamma_B \xrightarrow{\mathsf{q}:\mathsf{p}\&\ell_0}$ but we don't have $\Gamma_B \xrightarrow{(\mathsf{p},\mathsf{q})\ell_0}$ as $\mathtt{int} \not\leq \mathtt{nat}$.

Let $\Gamma_C = \mathsf{p} : \mathsf{q}{\oplus}\{\ell_1(\mathtt{int}).\mathsf{q}{\oplus}\{\ell_0(\mathtt{int}).\mathsf{end}\}\}, \mathsf{q} : \mathsf{p}\&\{\ell_1(\mathtt{int}).\mathsf{p}\&\{\ell_0(\mathtt{nat}).\mathsf{end}\}\}$. Γ_C is not safe as we have $\Gamma_C \xrightarrow{(\mathsf{p},\mathsf{q})\ell_1} \Gamma_B$ and Γ_B is not safe.

Consider Γ from Example 4.5. All the reducts satisfy [S-&⊕], hence Γ is safe.

Being a coinductive property, safe can be expressed in Rocq using Paco:

```
Definition weak_safety (c: tctx ) ≜
∀ p q s s'  k k', tctxRE (lsend p q (Some s) k) c  →  tctxRE (lrecv q p (Some s') k') c  →
                              tctxRE (lcomm p q k) c.

Inductive safe (R: tctx  →  Prop): tctx  →  Prop ≜
  | safety_red :  ∀ c, weak_safety c  →  (∀ p q c' k,
    tctxR c (lcomm p q k) c'  →  R c')
     →   safe R c.

Definition safeC c ≜ paco1 safe bot1 c.
```

`weak_safety` corresponds [S-&⊕] where `tctxRE l c` is shorthand for `∃ c', tctxR c l c'`. In the inductive `safe`, the constructor `safety_red` corresponds to [S-→]. Then `safeC` is defined as the greatest fixed point of `safe`.

We have that local type contexts with associated global types are always safe.

▶ **Theorem 5.3** (Safety by Association). *If* `assoc gamma g` *then* `safeC gamma`.

5.2 Fairness and Liveness

We now focus our attention to fairness and liveness. In this paper we have defined LTS semantics on three types of constructs: sessions, local type contexts and global types. We will appropriately define liveness properties on all three of these systems, so it will be convenient to define a general notion of valid reduction paths (also known as *runs* or *executions* [2, 2.1.1]) along with a general statement of some Linear Temporal Logic [35] constructs.

We start by defining the general notion of a reduction path [2, Def. 2.6] using possibly infinite cosequences.

▶ **Definition 5.4** (Reduction Paths). *A finite reduction path is an alternating sequence of states and labels* $S_0\lambda_0S_1\lambda_1...S_n$ *such that* $S_i \xrightarrow{\lambda_i} S_{i+1}$ *for all* $0 \leq i < n$. *An infinite reduction path is an alternating sequence of states and labels* $S_0\lambda_0S_1\lambda_1...S_n$ *such that* $S_i \xrightarrow{\lambda_i} S_{i+1}$ *for all* $0 \leq i$.

We won't be distinguishing between finite and infinite reduction paths and refer to them both as just *(reduction) paths*. Note that the above definition is general for LTSs, by *state* we will be referring to local type contexts, global types or sessions, depending on our application.

In Rocq, we define reduction paths using possibly infinite cosequences of pairs of states (which will be `tctx`, `gtt` or `session` in this paper) and `option label`:

```
CoInductive coseq (A: Type): Type ≜
  | conil : coseq A
  | cocons: A → coseq A → coseq A.
Notation local_path ≜ (coseq (tctx*option label)).
Notation global_path ≜ (coseq (gtt*option label)).
Notation session_path ≜ (coseq (session*option label)).
```

Note the use of `option label`, where we employ `None` to represent transitions into the end of the list, `conil`. For example, $S_0 \xrightarrow{\lambda_0} S_1 \xrightarrow{\lambda_1} S_2$ would be represented in Rocq as `cocons (s_0, Some lambda_0) (cocons (s_1, Some lambda_1) (cocons (s_2,None) conil))`, and `cocons (s_1, Some lambda) conil` would not be considered a valid path.

Note that this definition doesn't require the transitions in the `coseq` to actually be valid. We achieve that using the coinductive predicate `valid_path_GC A:Type (V: A→ label → A→ Prop)`, where the parameter `V` is a *transition validity predicate*, capturing if a one-step transition is valid. We use different `V` based on our application, for example in the context of local type context reductions `V gamma gamma'`

That is, we only allow synchronised communications in a valid local type context reduction path.

We can now define fairness and liveness on paths. We first restate the definition of fairness and liveness for local type context paths from [46], and use that to motivate our use of more general LTL constructs.

▶ **Definition 5.5** (Fair, Live Paths). *We say that a local type context path $\Gamma_0 \xrightarrow{\lambda_0} \Gamma_1 \xrightarrow{\lambda_2} ..$ is fair if, for all $n \in N : \Gamma_n \xrightarrow{(\mathsf{p},\mathsf{q})\ell}$ implies $\exists k, \ell'$ such that $N \ni k \geq n$ and $\lambda_k = (\mathsf{p},\mathsf{q})\ell'$, and therefore $\Gamma_k \xrightarrow{(\mathsf{p},\mathsf{q})\ell'} \Gamma_{k+1}$. We say that a path $(\Gamma_n)_{n \in N}$ is live iff, $\forall n \in N$:*

1. $\forall n \in N : \Gamma_n \xrightarrow{\mathsf{p}:\mathsf{q}\oplus\ell(S)}$ *implies $\exists k, \ell'$ such that $N \ni k \geq n$ and $\Gamma_k \xrightarrow{(\mathsf{p},\mathsf{q})\ell'} \Gamma_{k+1}$*
2. $\forall n \in N : \Gamma_n \xrightarrow{\mathsf{q}:\mathsf{p}\&\ell(S)}$ *implies $\exists k, \ell'$ such that $N \ni k \geq n$ and $\Gamma_k \xrightarrow{(\mathsf{p},\mathsf{q})\ell'} \Gamma_{k+1}$*

▶ **Definition 5.6** (Live Local Type Context). *A local type context Γ is live if whenever $\Gamma \to^* \Gamma'$, every fair path starting from Γ' is also live.*

In general, fairness assumptions are used so that only the reduction sequences that are "well-behaved" in some sense are considered when formulating other properties [44]. For our purposes we define fairness such that, in a fair path, if at any point p attempts to send to q *and* q attempts to send to p then eventually a communication between p and q takes place. Then live paths are defined to be paths such that whenever p attempts to send to q *or* q attempts to send to p, eventually a p to q communication takes place. Informally, this means that every communication request is eventually answered. Then live typing contexts are defined to be the Γ where all fair paths that start from Γ are also live.

redo

▶ **Example 5.7.** Consider the contexts Γ, Γ' and Γ_{end} from Example 4.5. One possible reduction path is $\Gamma \xrightarrow{(\mathsf{p},\mathsf{q})\ell_0} \Gamma \xrightarrow{(\mathsf{p},\mathsf{q})\ell_0} \ldots$. Denote this path as $(\Gamma_n)_{n \in \mathbb{N}}$, where $\Gamma_n = \Gamma$ for all $n \in \mathbb{N}$. By reductions (??) and (??), we have $\forall n, \Gamma_n \xrightarrow{(\mathsf{p},\mathsf{q})\ell_0}$ and $\Gamma_n \xrightarrow{(\mathsf{p},\mathsf{q})\ell_1}$ as the only possible synchronised reductions from Γ_n. Accordingly, we also have $\forall n, \Gamma_n \xrightarrow{(\mathsf{p},\mathsf{q})\ell_0} \Gamma_{n+1}$ in the path so this path is fair. However, this path is not live as we have by reduction (??) that $\Gamma_1 \xrightarrow{\mathsf{r}:\mathsf{q}\&\ell_2(\mathsf{int})}$ but there is no n, ℓ' with $\Gamma_n \xrightarrow{(\mathsf{q},\mathsf{r})\ell'} \Gamma_{n+1}$ in the path. Consequently, Γ is not a live type context.

Now consider the reduction path $\Gamma \xrightarrow{(\mathsf{p},\mathsf{q})\ell_0} \Gamma \xrightarrow{(\mathsf{p},\mathsf{q})\ell_0} \Gamma' \xrightarrow{(\mathsf{q},\mathsf{r})\ell_2} \Gamma_{\mathtt{end}}$, denoted by $(\Gamma'_n)_{n\in\{1..4\}}$. This path is fair with respect to reductions from Γ'_1 and Γ'_2 as shown above, and it's fair with respect to reductions from Γ'_3 as reduction (**??**) is the only one available from Γ'_3 and we have $\Gamma'_3 \xrightarrow{(\mathsf{q},\mathsf{r})\ell_2} \Gamma'_4$ as needed. Furthermore, this path is live: the reduction $\Gamma_1 \xrightarrow{\mathsf{r}:\mathsf{q}\&\ell_2(\mathtt{int})}$ that causes (Γ_n) to fail liveness is handled by the reduction $\Gamma'_3 \xrightarrow{(\mathsf{q},\mathsf{r})\ell_2} \Gamma'_4$ in this case.

Definition 5.5 , while intuitive, is not really convenient for a Rocq formalisation due to the existential statements contained in them. It would be ideal if these properties could be expressed as a least or greatest fixed point, which could then be formalised via Rocq's inductive or coinductive (via Paco) types. To achieve this, we recast fairness and liveness for local type context paths in Linear Temporal Logic (LTL) [35]. Specifically, define atomic propositions $\mathtt{enabledComm}_{\mathsf{p},\mathsf{q},\ell}$ such that $(\Gamma,\lambda) \models \mathtt{enabledComm}_{\mathsf{p},\mathsf{q},\ell} \iff \Gamma \xrightarrow{(\mathsf{p},\mathsf{q})\ell}$, and $\mathtt{headComm}_{\mathsf{p},\mathsf{q}}$ that holds iff $\lambda = (\mathsf{p},\mathsf{q})\ell$ for some ℓ. Then fairness can be expressed in LTL with: for all p,q,

$$\Box(\mathtt{enabledComm}_{\mathsf{p},\mathsf{q},\ell} \implies \Diamond(\mathtt{headComm}_{\mathsf{p},\mathsf{q}}))$$

Similarly, by defining $\mathtt{enabledSend}_{\mathsf{p},\mathsf{q},\ell,S}$ that holds iff $\Gamma \xrightarrow{\mathsf{p}:\mathsf{q}\oplus\ell(S)}$ and analogously $\mathtt{enabledRecv}$, liveness can be defined as

$$\begin{aligned}&\Box((\mathtt{enabledSend}_{\mathsf{p},\mathsf{q},\ell,S} \implies \Diamond(\mathtt{headComm}_{\mathsf{p},\mathsf{q}}))\wedge\\ &\quad(\mathtt{enabledRecv}_{\mathsf{p},\mathsf{q},\ell,S} \implies \Diamond(\mathtt{headComm}_{\mathsf{q},\mathsf{p}})))\end{aligned}$$

The reason we defined the properties using LTL properties is that the operators $\Diamond$ and $\Box$ can be characterised as least and greatest fixed points using their expansion laws [2, Chapter 5.14]:

- $\Diamond P$ is the least solution to $\Diamond P \equiv P \vee \bigcirc(\Diamond P)$
- $\Box P$ is the greatest solution to $\Box P \equiv P \wedge \bigcirc(\Box P)$

Thus fairness and liveness correspond to greatest fixed points, which can be defined coinductively.

In Rocq, we implement the LTL operators $\Diamond$ and $\Box$ inductively and coinductively (with Paco), in the following way:

```
Inductive eventually {A: Type} (F: coseq A → Prop): coseq A → Prop ≜
  | evh: ∀ xs, F xs → eventually F xs
  | evc: ∀ x xs, eventually F xs → eventually F (cocons x xs).

Inductive until {A:Type} (F: coseq A → Prop) (G: coseq A → Prop) : coseq A → Prop≜
  | untilh : ∀ xs, G xs → until F G xs
  | untilc: ∀ x xs, F (cocons x xs) → until F G xs → until F G (cocons x xs).

Inductive alwaysG {A: Type} (F: coseq A → Prop) (R: coseq A → Prop): coseq A → Prop ≜
  | alwn: F conil → alwaysG F R conil
  | alwc: ∀ x xs, F (cocons x xs) → R xs → alwaysG F R (cocons x xs).

Definition alwaysCG {A:Type} (F: coseq A → Prop) ≜ paco1 (alwaysG F) bot1.
```

Note the use of the constructor `alwn` in the definition `alwaysG` to handle finite paths.

Using these LTL constructs we can define fairness and liveness on paths.

```
Definition fair_path_local_inner (pt: local_path): Prop ≜
  ∀ p q n, to_path_prop (tctxRE (lcomm p q n)) False pt →  eventually (headComm p q) pt.
Definition fair_path ≜ alwaysCG fair_path_local_inner.
Definition live_path_inner (pt: local_path) : Prop ≜ ∀ p q s n,
(to_path_prop (tctxRE (lsend p q (Some s) n)) False pt → eventually (headComm p q) pt) ∧
(to_path_prop (tctxRE (lrecv p q (Some s) n)) False pt → eventually (headComm q p) pt).
Definition live_path ≜ alwaysCG live_path_inner.
```

For instance, the fairness of the first reduction path for Γ given in Example 5.7 can be expressed with the following:

```
CoFixpoint inf_pq_path ≜ cocons (gamma,(lcomm prt_p prt_q) 0) inf_pq_path.
Theorem inf_pq_path_fair : fairness inf_pq_path.
```

▶ Remark 5.8. Note that the LTS of local type contexts has the property that, once a transition between participants p and q is enabled, it stays enabled until a transition between p and q occurs. This makes `fair_path` equivalent to the standard formulas [2, Definition 5.25] for strong fairness ($\Box\Diamond\mathtt{enabledComm}_{\mathsf{p},\mathsf{q}} \implies \Box\Diamond\mathtt{headComm}_{\mathsf{p},\mathsf{q}}$) and weak fairness ($\Diamond\Box\mathtt{enabledComm}_{\mathsf{p},\mathsf{q}} \implies \Box\Diamond\mathtt{headComm}_{\mathsf{p},\mathsf{q}}$).

With these definitions we can now prove that local type contexts associated with a global type are live, which is the most involved of the results mechanised in this work. We now detail the Rocq Proof that associated local type contexts are also live.

▶ Remark 5.9. We once again emphasise that all global types mentioned are assumed to be balanced (Definition 3.10). Indeed association with non-balanced global types doesn't guarantee liveness. As an example, consider Γ from Example 4.5, which is associated with G from Example 4.8. Yet we have shown in Example 5.7 that Γ is not a live type context. This is not surprising as G is not balanced.

▶ **Theorem 5.10** (Liveness by Association). *If* `assoc gamma g` *then* `gamma` *is live.*

Proof. (Simplified, Outline) Our proof proceeds in two steps. First, we prove that the typing context obtained by direct projections of `g`, that is, `gamma_proj` = $\{\mathsf{p}_i : \mathsf{G}\restriction_{\mathsf{p}_i} \mid \mathsf{p}_i \in \mathtt{pt}\{\mathsf{G}\}\}$, is live. We then leverage Theorem 4.10 to show that if `gamma_proj` is live, so is `gamma`.

The proof that `gamma_proj` is live proceeds by well-founded induction on the tree height [13] of the grafting (Lemma 3.13) of the global type `g`. Suppose `gamma_proj` $\xrightarrow{\mathsf{p}:\mathsf{q}\oplus\ell(S)}$ (the case for the receive is similar and omitted), and `xs` is a fair local type context reduction path beginning with `gamma_proj`. To show that `xs` is live we need to show the existence of a $(\mathsf{p},\mathsf{q})\ell$ transition in `xs`. We prove the following helper lemmas:

- The height of the `p`-grafting of `g` is not smaller than the `q`-grafting.
- If the `p`-grafting and `q`-grafting of a global type `g'` have the same height, then any fair path beginning with the direct projection context of `g'` eventually contains a $(\mathsf{p},\mathsf{q})\ell$ transition.
- The height of the `p`-grafting of `g` strictly decreases with every transition involving `q`, and doesn't increase with the transitions not involving `q`.

These lemmas followed by well-founded induction on the height of the `p`-grafting of the global type the head of `xs` is projected from gives the desired transition.

In the second step of the proof we extend association on to paths to get, for each local type context reduction path `xs` that begins with `gamma`, another local reduction path `ys` beginning with `gamma_proj` such that the elements of `xs` are subtypes (subtyping on contexts defined pointwise) of the corresponding elements of `ys`. This is obtained from Theorem 4.10, however the statement of Theorem 4.10 is implemented as an $\exists$ statement that lives in `Prop`, hence we need to use the `constructive_indefinite_description` axiom to construct a `CoFixpoint` returning the desired cosequence `ys`. The proof then follows by the definition of subtyping (Definition 3.5).

Suppose `gamma` $\to^*$ `gamma'`, then by Theorem 4.10 `assoc gamma' g'` for some `g'`, and hence by Corollary 5.23 any fair path starting from `gamma'` is live, as needed. ◀

Our proof proceeds in the following way:

1. Formulate an analogue of fairness and liveness for global type reduction paths.
2. Prove that all global types are live for this notion of liveness.
3. Show that if `G : gtt` is live and `assoc gamma G`, then `gamma` is also live.

First we define fairness and liveness for global types, analogous to Definition 5.5.

▶ **Definition 5.11** (Fairness and Liveness for Global Types). *We say that the label λ is enabled at G if the context $\{\mathsf{p}_i : \mathsf{G} \upharpoonright_{\mathsf{p}_i} \;\;\|\;\; \mathsf{p}_i \in \mathtt{pt}\{\mathsf{G}\}\}$ can transition via λ. More explicitly, and in Rocq terms,*

```
Definition global_label_enabled l g≜ match l with
  | lsend p q (Some s) n ⇒ ∃ xs g',
    projectionC g p  (ltt_send q xs) ∧ onth n xs=Some (s,g')
  | lrecv p q (Some s) n ⇒ ∃ xs g',
    projectionC g p  (ltt_recv q xs) ∧ onth n xs=Some (s,g')
  | lcomm p q n ⇒ ∃ g', gttstepC g g' p q n
  | _ ⇒ False end.
```

With this definition of enabling, fairness and liveness are defined exactly as in Definition 5.5. A global type reduction path is fair if the following holds:

$$\Box(\mathtt{enabledComm}_{\mathsf{p},\mathsf{q},\ell} \implies \Diamond(\mathtt{headComm}_{\mathsf{p},\mathsf{q}}))$$

and liveness is expressed with the following:

$$\Box((\mathtt{enabledSend}_{\mathsf{p},\mathsf{q},\ell,S} \implies \Diamond(\mathtt{headComm}_{\mathsf{p},\mathsf{q}}))\wedge$$
$$(\mathtt{enabledRecv}_{\mathsf{p},\mathsf{q},\ell,S} \implies \Diamond(\mathtt{headComm}_{\mathsf{q},\mathsf{p}})))$$

where `enabledSend`, `enabledRecv` *and* `enabledComm` *correspond to the match arms in the definition of* `global_label_enabled` *(Note that the names* `enabledSend` *and* `enabledRecv` *are chosen for consistency with Definition 5.5, there aren't actually any transitions with label $\mathsf{p} : \mathsf{q} \oplus \ell(S)$ in the transition system for global types). A global type G is live if whenever $\mathsf{G} \to^* \mathsf{G}'$, any fair path starting from G' is also live.*

Now our goal is to prove that all (well-formed, balanced, projectable) G are live under this definition. This is where the notion of grafting (Definition 3.10) becomes important, as the proof essentially proceeds by well-founded induction on the height of the tree obtained by grafting.

We first introduce some definitions on global type tree contexts (Definition 3.11).

▶ **Definition 5.12** (Global Type Context Equality, Proper Prefixes and Height). *We consider two global type tree contexts to be equal if they are the same up to the relabelling the indices of their leaves. More precisely,*

```
Inductive gtth_eq: gtth → gtth → Prop ≜
  | gtth_eq_hol : ∀ n m, gtth_eq (gtth_hol n) (gtth_hol m)
  | gtth_eq_send : ∀ xs ys p q ,
    Forall2 (fun u v ⇒ (u=None ∧ v=None) ∨ (∃ s g1 g2, u=Some (s,g1) ∧ v=Some (s,g2) ∧ gtth_eq g1 g2)) xs ys →
    gtth_eq (gtth_send p q xs) (gtth_send p q ys).
```

Informally, we say that the global type context $\mathbb{G}'$ is a proper prefix of $\mathbb{G}$ if we can obtain $\mathbb{G}'$ by changing some subtrees of $\mathbb{G}$ with context holes such that none of the holes in $\mathbb{G}$ are present in $\mathbb{G}'$. Alternatively, we can characterise it as akin to `gtth_eq` *except where the context holes in $\mathbb{G}'$ are assumed to be "jokers" that can be matched with any global type context that's not just a context hole. In Rocq:*

```
Inductive is_tree_proper_prefix : gtth → gtth → Prop ≜
  | tree_proper_prefix_hole : ∀ n p q xs, is_tree_proper_prefix (gtth_hol n) (gtth_send p q xs)
  | tree_proper_prefix_tree : ∀ p q xs ys,
    Forall2 (fun u v ⇒ (u=None ∧ v=None)
      ∨ ∃ s g1 g2, u=Some (s, g1) ∧ v=Some (s,g2) ∧
              is_tree_proper_prefix g1 g2
    ) xs ys →
    is_tree_proper_prefix (gtth_send p q xs) (gtth_send p q ys).
```

give examples

We also define a function `gtth_height : gtth → Nat` *that computes the height [13] of a global type tree context. Context holes i.e. leaves have height 0, and the height of an internal node is the maximum of the height of their children plus one.*

```
Fixpoint gtth_height (gh : gtth) : nat ≜
  match gh with
  | gtth_hol n ⇒ 0
  | gtth_send p q xs ⇒
    list_max (map (fun u⇒ match u with
         | None ⇒ 0
         | Some (s,x) ⇒ gtth_height x end) xs) + 1 end.
```

`gtth_height`, `gtth_eq` and `is_tree_proper_prefix` interact in the expected way.

▶ **Lemma 5.13.** *If* `gtth_eq gx gx'` *then* `gtth_height gx = gtth_height gx'`*.*

▶ **Lemma 5.14.** *If* `is_tree_proper_prefix gx gx'` *then* `gtth_height gx < gtth_height gx'`*.*

Our motivation for introducing these constructs on global type tree contexts is the following *multigrafting* lemma:

▶ **Lemma 5.15** (Multigrafting)**.** *Let* `projectionC g p (ltt_send q xsp)` *or* `projectionC g p (ltt_recv q xsp)`*,* `projectionC g q Tq`*,* `g` *is* `p`*-grafted by* `ctx_p` *and* `gs_p`*, and* `g` *is* `q`*-grafted by* `ctx_q` *and* `gs_q`*. Then either* `is_tree_proper_prefix ctx_q ctx_p` *or* `gtth_eq ctx_p ctx_q`*. Furthermore, if* `gtth_eq ctx_p ctx_q` *then* `projectionC g q (ltt_send p xsq)` *or* `projectionC g q (ltt_recv p xsq)` *for some* `xsq`*.*

Proof. By induction on the global type context `ctx_p`. ◀

example

We also have that global type reductions that don't involve participant `p` can't increase the height of the `p`-grafting, established by the following lemma:

▶ **Lemma 5.16.** *Suppose* `g : gtt` *is* `p`*-grafted by* `gx : gtth` *and gs : list (option gtt),* `gttstepC g g' s t ell` *where* `p` $\neq$ `s` *and* `p` $\neq$ `t`*, and* `g'` *is* `p`*-grafted by* `gx'` *and* `gs'`*. Then*
- **(i)** *If* `ishParts s gx` *or* `ishParts t gx` *, then* `gtth_height gx' < gtth_height gx`
- **(ii)** *In general,* `gtth_height gx'` $\leq$ `gtth_height gx`

Proof. We define a inductive predicate `gttstepH : gtth → part → part → part → gtth → Prop` with the property that if `gttstepC g g' p q ell` for some `r` $\neq$ `p, q`, and tree contexts `gx` and `gx'` `r`-graft `g` and `g'` respectively, then `gttstepH gx p q ell gx'` (`gttstepH_consistent`). The results then follow by induction on the relation `gttstepH gx s t ell gx'`. ◀

We can now prove the liveness of global types. The bulk of the work goes in to proving the following lemma:

▶ **Lemma 5.17.** *Let* `xs` *be a fair global type reduction path starting with* `g`*.*
- **(i)** *If* `projectionC g p (ltt_send q xsp)` *for some* `xsp`*, then a* `lcomm p q ell` *transition takes place in* `xs` *for some message label* `ell`*.*

(ii) *If* `projectionC g p (ltt_recv q xsp)` *for some* `xsp`*, then a* `lcomm q p ell` *transition takes place in* `xs` *for some message label* `ell`*.*

Proof. We outline the proof for (i), the case for (ii) is symmetric.

Rephrasing slightly, we prove the following: forall `n : nat` and global type reduction path `xs`, if the head `g` of `xs` is `p`-grafted by `ctx_p` and `gtth_height ctx_p = n`, the lemma holds. We proceed by strong induction on `n`, that is, the tree context height of `ctx_p`.

Let `(ctx_q, gs_q)` be the `q`-grafting of `g`. By Lemma 5.15 we have that either `gtth_eq ctx_q ctx_p` (a) or `is_tree_proper_prefix ctx_q ctx_p` (b). In case (a), we have that `projectionC g q (ltt_recv p xsq)`, hence by (cite simul subproj or something here) and fairness of `xs`, we have that a `lcomm p q ell` transition eventually occurs in `xs`, as required.

In case (b), by Lemma 5.14 we have `gtth_height ctx_q < gtth_height ctx_p`, so by the induction hypothesis a transition involving `q` eventually happens in `xs`. Assume wlog that this transition has label `lcomm q r ell`, or, in the pen-and-paper notation, $(\mathsf{q},\mathsf{r})\ell$. Now consider the prefix of `xs` where the transition happens: `g` $\xrightarrow{\lambda}$ `g_1` $\rightarrow$ .. `g'` $\xrightarrow{(\mathsf{q},\mathsf{r})\ell}$ `g''`. Let `g'` be `p`-grafted by the global tree context `ctx'_p`, and `g''` by `ctx''_p`. By Lemma 5.16, `gtth_height ctx''_p < gtth_height ctx'_p` $\leq$ `gtth_height ctx_p`. Then, by the induction hypothesis, the suffix of `xs` starting with `g''` must eventually have a transition `lcomm p q ell'` for some `ell'`, therefore `xs` eventually has the desired transition too. ◀

Lemma 5.17 proves that any fair global type reduction path is also a live path, from which the liveness of global types immediately follows.

▶ **Corollary 5.18.** *All global types are live.*

We can now leverage the simulation established by Theorem 4.10 to prove the liveness (Definition 5.5) of local typing context reduction paths.

We start by lifting association (Definition 4.7) to reduction paths.

▶ **Definition 5.19** (Path Association)**.** *Path association is defined coinductively by the following rules:*

(i) *The empty path is associated with the empty path.*

(ii) *If* $\Gamma \xrightarrow{\lambda_0} \rho$ *is path-associated with* $\mathsf{G} \xrightarrow{\lambda_1} \rho'$ *where (*ρ *and* ρ' *are local and global reduction paths, respectively), then* $\lambda_0 = \lambda_1$ *and* ρ *is path-associated with* ρ'*.*

```
Variant path_assoc (R:local_path → global_path → Prop): local_path → global_path → Prop ≜
  | path_assoc_nil : path_assoc R conil conil
  | path_assoc_xs : ∀ g gamma l xs ys, assoc gamma g → R xs ys →
  path_assoc R (cocons (gamma, l) xs) (cocons (g, l) ys).

Definition path_assocC ≜ paco2 path_assoc bot2.
```

Informally, a local type context reduction path is path-associated with a global type reduction path if their matching elements are associated and have the same transition labels.

We show that reduction paths starting with associated local types can be path-associated.

▶ **Lemma 5.20.** *If* `assoc gamma g`*, then any local type context reduction path starting with* `gamma` *is associated with a global type reduction path starting with* `g`*.*

Proof. Let the local reduction path be `gamma` $\xrightarrow{\lambda}$ `gamma_1` $\xrightarrow{\lambda_1}$ .... We construct a path-associated global reduction path. By Theorem 4.10 there is a `g_1 : gtt` such that `g` $\xrightarrow{\lambda}$ `g_1` and `assoc gamma_1 g_1`, hence the path-associated global type reduction path starts with `g`

maybe just give the definition as a cofixpoint?

$\xrightarrow{\lambda}$ g_1. We can repeat this procedure to the remaining path starting with gamma_1 $\xrightarrow{\lambda_1}$ ... to get g_2 : gtt such that assoc gamma_2 g_2 and g_1 $\xrightarrow{\lambda_1}$ g_2. Repeating this, we get g $\xrightarrow{\lambda}$ g_1 $\xrightarrow{\lambda_1}$ .. as the desired path associated with gamma $\xrightarrow{\lambda}$ gamma_1 $\xrightarrow{\lambda_1}$ .... ◀

▶ Remark 5.21. In the Rocq implementation the construction above is implemented as a CoFixpoint returning a coseq. Theorem 4.10 is implemented as an ∃ statement that lives in Prop, hence we need to use the constructive_indefinite_description axiom to obtain the witness to be used in the construction.

We also have the following correspondence between fairness and liveness properties for associated global and local reduction paths.

▶ **Lemma 5.22.** *For a local reduction path* xs *and global reduction path* ys*, if* path_assocC xs ys *then*

(i) *If* xs *is fair then so is* ys

(ii) *If* ys *is live then so is* xs

As a corollary of Lemma 5.22, Lemma 5.20 and Lemma 5.17 we have the following:

▶ **Corollary 5.23.** *If* assoc gamma g*, then any fair local reduction path starting from* gamma *is live.*

Proof. Let xs be the fair local reduction path starting with gamma. By Lemma 5.20 there is a global path ys associated with it. By Lemma 5.22 (i) ys is fair, and by Lemma 5.17 ys is live, so by Lemma 5.22 (ii) xs is also live. ◀

Liveness of contexts follows directly from Corollary 5.23.

▶ **Theorem 5.24** (Liveness by Association)**.** *If* assoc gamma g *then* gamma *is live.*

Proof. Suppose gamma $\rightarrow^*$ gamma', then by Theorem 4.10 assoc gamma' g' for some g', and hence by Corollary 5.23 any fair path starting from gamma' is live, as needed. ◀

6 Properties of Sessions

We give typing rules for the session calculus introduced in 2, and prove subject reduction and progress for them. Then we define a liveness property for sessions, and show that processes typable by a local type context that's associated with a global type tree are guaranteed to satisfy this liveness property.

6.1 Typing rules

We give typing rules for our session calculus based on [18] and [15].

We distinguish between two kinds of typing judgements and type contexts.

1. A local type context Γ associates participants with local type trees, as defined in cdef-type-ctx. Local type contexts are used to type sessions (Definition 2.2) i.e. a set of pairs of participants and single processes composed in parallel. We express such judgements as $\Gamma \vdash_{\mathcal{M}} \mathcal{M}$, or as typ_sess M gamma or gamma ⊢ M in Rocq.
2. A process variable context Θ_T associates process variables with local type trees, and an expression variable context Θ_e assigns sorts to expresion variables. Variable contexts are used to type single processes and expressions (Definition 2.1). Such judgements are expressed as $\Theta_\mathsf{T}, \Theta_\mathsf{e} \vdash_\mathsf{P} \mathsf{P} : \mathsf{T}$, or in Rocq as typ_proc theta_T theta_e P T or theta_T, theta_e ⊢ P : T.

$$\Theta \vdash_{\mathsf{P}} n\colon \mathtt{nat} \qquad \Theta \vdash_{\mathsf{P}} i\colon \mathtt{int} \qquad \Theta \vdash_{\mathsf{P}} \mathtt{true}\colon \mathtt{bool} \qquad \Theta \vdash_{\mathsf{P}} \mathtt{false}\colon \mathtt{bool} \qquad \Theta, x\colon \mathsf{S} \vdash_{\mathsf{P}} x\colon \mathsf{S}$$

$$\frac{\Theta \vdash_{\mathsf{P}} e\colon \mathtt{nat}}{\Theta \vdash_{\mathsf{P}} \mathtt{succ}\ e\colon \mathtt{nat}} \qquad \frac{\Theta \vdash_{\mathsf{P}} e\colon \mathtt{int}}{\Theta \vdash_{\mathsf{P}} \mathtt{neg}\ e\colon \mathtt{int}} \qquad \frac{\Theta \vdash_{\mathsf{P}} e\colon \mathtt{bool}}{\Theta \vdash_{\mathsf{P}} \neg\ e\colon \mathtt{bool}}$$

$$\frac{\Theta \vdash_{\mathsf{P}} e_1\colon \mathsf{S} \quad \Theta \vdash_{\mathsf{P}} e_2\colon \mathsf{S}}{\Theta \vdash_{\mathsf{P}} e_1 \oplus e_2\colon \mathsf{S}} \qquad \frac{\Theta \vdash_{\mathsf{P}} e_1\colon \mathtt{int} \quad \Theta \vdash_{\mathsf{P}} e_2\colon \mathtt{int}}{\Theta \vdash_{\mathsf{P}} e_1 > e_2\colon \mathtt{bool}} \qquad \frac{\Theta \vdash_{\mathsf{P}} e\colon \mathsf{S} \quad \mathsf{S} \leq \mathsf{S}'}{\Theta \vdash_{\mathsf{P}} e\colon \mathsf{S}'}$$

Table 4 Typing expressions

$$[\text{T-END}]\ \Theta \vdash_{\mathsf{P}} \mathbf{0}\colon \mathtt{end} \qquad [\text{T-VAR}]\ \Theta, \mathbf{X}\colon \mathsf{T} \vdash_{\mathsf{P}} \mathbf{X}\colon \mathsf{T} \qquad [\text{T-REC}]\ \frac{\Theta, \mathbf{X}\colon \mathsf{T} \vdash_{\mathsf{P}} \mathsf{P}\colon \mathsf{T}}{\Theta \vdash_{\mathsf{P}} \mu\mathbf{X}.\mathsf{P}\colon \mathsf{T}} \qquad [\text{T-IF}]\ \frac{\Theta \vdash_{\mathsf{P}} e\colon \mathtt{bool} \quad \Theta \vdash_{\mathsf{P}} \mathsf{P}_1\colon \mathsf{T} \quad \Theta \vdash_{\mathsf{P}} \mathsf{P}_2\colon \mathsf{T}}{\Theta \vdash_{\mathsf{P}} \mathtt{if}\ \mathsf{e}\ \mathtt{then}\ \mathsf{P}_1\ \mathtt{else}\ \mathsf{P}_2\colon \mathsf{T}}$$

$$[\text{T-SUB}]\ \frac{\Theta \vdash_{\mathsf{P}} \mathsf{P}\colon \mathsf{T} \quad \mathsf{T} \leqslant \mathsf{T}'}{\Theta \vdash_{\mathsf{P}} \mathsf{P}\colon \mathsf{T}'} \qquad [\text{T-IN}]\ \frac{\forall i \in I, \quad \Theta, x_i\colon \mathsf{S}_i \vdash_{\mathsf{P}} \mathsf{P}_i\colon \mathsf{T}_i}{\Theta \vdash_{\mathsf{P}} \sum_{i \in I} \mathsf{p}?\ell_i(x_i).\mathsf{P}_i\colon \mathsf{p}\&\{\ell_i(\mathsf{S}_i).\mathsf{T}_i\}_{i \in I}} \qquad [\text{T-OUT}]\ \frac{\Theta \vdash_{\mathsf{P}} e\colon \mathsf{S} \quad \Theta \vdash_{\mathsf{P}} \mathsf{P}\colon \mathsf{T}}{\Theta \vdash_{\mathsf{P}} \mathsf{p}!\ell(\mathsf{e}).\mathsf{P}\ :\ \mathsf{p}\oplus\{\ell(\mathsf{S}).\mathsf{T}\}}$$

Table 5 Typing processes

Table 4 and Table 5 state the standard typing rules for expressions and processes which we don't elaborate on. We have a single rule for typing sessions:

$$[\text{T-SESS}]\ \frac{\forall i \in I: \quad \vdash_{\mathsf{P}} \mathsf{P}_i\colon \Gamma(\mathsf{p}_i) \qquad \Gamma \sqsubseteq \mathsf{G}}{\Gamma \vdash_{\mathcal{M}} \prod_i \mathsf{p}_i \triangleleft \mathsf{P}_i}$$

[T-SESS] says that a session made of the parallel composition of processes $\prod_i \mathsf{p}_i \triangleleft \mathsf{P}_i$ can be typed by an associated local context Γ if the local type of participant p_i in Γ types the process

6.2 Subject Reduction, Progress and Session Fidelity

The subject reduction, progress and non-stuck theorems from [15] also hold in this setting, with minor changes in their statements and proofs. We won't discuss these proofs in detail.

give theorem no

▶ **Lemma 6.1.** *If* `gamma` $\vdash_{\mathcal{M}}$ `M` *and* `M` $\Rrightarrow$ `M'` *then* `typ_sess M' gamma`.

▶ **Theorem 6.2** (Subject Reduction). *If* `gamma` $\vdash_{\mathcal{M}}$ `M` *and* `M` $\xrightarrow{(\mathsf{p},\mathsf{q})\ell}$ `M'`, *then there exists a typing context* `gamma'` *such that* `gamma` $\xrightarrow{(\mathsf{p},\mathsf{q})\ell}$ `gamma'` *and* `gamma` $\vdash_{\mathcal{M}}$ `M` .

▶ **Theorem 6.3** (Progress). *If* `gamma` $\vdash_{\mathcal{M}}$ `M` , *one of the following hold :*

1. *Either* `M` $\Rrightarrow$ `M_inact` *where every process making up* `M_inact` *is inactive, i.e.* `M_inact` $\equiv \prod_{i=1}^{n} \mathsf{p}_i \triangleleft \mathbf{0}$ *for some n.*
2. *Or there is a* `M'` *such that* `M` $\to$ `M'`.

▶ Remark 6.4. Note that in Theorem 6.2 one transition between sessions corresponds to exactly one transition between local type contexts with the same label. That is, every session transition is observed by the corresponding type. This is the main reason for our choice of reactive semantics (Section 2.2) as τ transitions are not observed by the type in ordinary semantics. In other words, with τ-semantics the typing relation is a *weak simulation* [30], while it turns into a strong simulation with reactive semantics. For our Rocq implementation working with the strong simulation turns out be more convenient.

We can also prove the following correspondence result in the reverse direction to Theorem 6.2, analogus to Theorem 4.9.

▶ **Theorem 6.5** (Session Fidelity). *If* `gamma` $\vdash_{\mathcal{M}}$ `M` *and* `gamma` $\xrightarrow{(\mathsf{p},\mathsf{q})\ell}$ `gamma'`*, there exists a message label* ℓ'*, a context* `gamma''` *and a session* `M'` *such that* `M` $\xrightarrow{(\mathsf{p},\mathsf{q})\ell'}$ `M'` , `gamma` $\xrightarrow{(\mathsf{p},\mathsf{q})\ell'}$ `gamma''` *and* `typ_sess M' gamma''`*.*

Proof. By inverting the local type context transition and the typing. ◀

▶ Remark 6.6. Again we note that by Theorem 6.5 a single-step context reduction induces a single-step session reduction on the type. With the τ-semantics the session reduction induced by the context reduction would be multistep.

Now the following type safety property follows from the above theorems:

▶ **Theorem 6.7** (Type Safety). *If* `gamma` $\vdash_{\mathcal{M}}$ `M` *and* `M` $\rightarrow^*$ `M'` $\Rrightarrow$ `p` $\leftarrow$ `p_send q ell P ||| q` $\leftarrow$ `p_recv p xs ||| M''`*, then* `onth ell xs` $\neq$ `None`*.*

6.3 Session Liveness

We state the liveness property we are interested in proving, and show that typable sessions have this property.

▶ **Definition 6.8** (Session Liveness). *Session* $\mathcal{M}$ *is live iff*

1. $\mathcal{M} \longrightarrow^* \mathcal{M}' \Rrightarrow \mathsf{q} \triangleleft \mathsf{p}!\ell_i(x_i).Q \mid \mathcal{N}$ *implies* $\mathcal{M}' \longrightarrow^* \mathcal{M}'' \Rrightarrow \mathsf{q} \triangleleft Q \mid \mathcal{N}'$ *for some* $\mathcal{M}'',\mathcal{N}'$
2. $\mathcal{M} \longrightarrow^* \mathcal{M}' \Rrightarrow \mathsf{q} \triangleleft \bigwedge_{i\in I} \mathsf{p}?\ell_i(x_i).Q_i \mid \mathcal{N}$ *implies* $\mathcal{M}' \longrightarrow^* \mathcal{M}'' \Rrightarrow \mathsf{q} \triangleleft Q_i[v/x_i] \mid \mathcal{N}'$ *for some* $\mathcal{M}'',\mathcal{N}', i, v$.

In Rocq we express this with the following:

```
Definition live_sess Mp ≜ ∀ M, betaRtc Mp M →
  (∀ p q ell e P' M', p ≠q → unfoldP M ( (p ←- p_send q ell e P') \|\|\| M') → ∃ M'',
  betaRtc M ((p ←- P')\|\|\|M''))
  ∧
  (∀ p  q llp M', p ≠q → unfoldP M ( (p ←- p_recv q llp) \|\|\| M') →
    ∃ M'' P' e k, onth k llp = Some P' ∧ betaRtc M ((p ←- subst_expr_proc P' e 0 0)\|\|\|M'')).
```

Session liveness, analogous to liveness for typing contexts (Definition 5.5), says that when $\mathcal{M}$ is live, if $\mathcal{M}$ reduces to a session $\mathcal{M}'$ containing a participant that's attempting to send or receive, then $\mathcal{M}'$ reduces to a session where that communication has happened. It's also called *lock-freedom* in related work ([43, 31]).

We now prove that typed sessions are live. Our proof follows the following steps:

1. Formulate a "fairness" property for typable sessions, with the property that any finite session reduction path can be extended to a fair session reduction path.
2. Lift the typing relation to reduction paths, and show that fair session reduction paths are typed by fair local type context reduction paths.
3. Prove that a certain transition eventually happens in the local context reduction path, and that this means the desired transition is enabled in the session reduction path.

We first state a "fairness" (the reason for the quotes is explained in Remark 6.10) property for session reduction paths, analogous to fairness for local type context reduction paths (Definition 5.5).

▶ **Definition 6.9** ("Fairness" of Sessions). *We say that a* $(\mathsf{p},\mathsf{q})\ell$ *transition is enabled at* $\mathcal{M}$ *if* $\mathcal{M} \xrightarrow{(\mathsf{p},\mathsf{q})\ell} \mathcal{M}'$ *for some* $\mathcal{M}'$*. A session reduction path is fair if the following LTL property holds:*

$$\square(\mathtt{enabledComm}_{\mathsf{p},\mathsf{q},\ell} \implies \lozenge(\mathtt{headComm}_{\mathsf{p},\mathsf{q}}))$$

▶ Remark 6.10. Definition 6.9 is not actually a sensible fairness property for our reactive semantics, mainly because it doesn't satisfy the *feasibility* [44] property stating that any finite execution can be extended to a fair execution. Consider the following session:

$$\mathcal{M} = \mathsf{p} \triangleleft \mathsf{if}(\mathsf{true} \oplus \mathsf{false})\ \mathsf{then}\ \mathsf{q}!\ell_1(\mathit{true})\ \mathsf{else}\ \mathsf{r}!\ell_2(\mathit{true}).\mathbf{0} \mid \mathsf{q} \triangleleft \mathsf{p}?\ell_\mathbf{1}(\mathbf{x}).\mathbf{0} \mid \mathsf{r} \triangleleft \mathsf{p}?\ell_\mathbf{2}(\mathbf{x}).\mathbf{0}$$

We have that $\mathcal{M} \xrightarrow{(\mathsf{p},\mathsf{q})\ell_1} \mathcal{M}'$ where $\mathcal{M}' = \mathsf{p} \triangleleft \mathbf{0} \mid \mathsf{q} \triangleleft \mathbf{0} \mid \mathsf{r} \triangleleft \mathsf{p}?\ell_\mathbf{2}(\mathbf{x}).\mathbf{0}$, and also $\mathcal{M} \xrightarrow{(\mathsf{p},\mathsf{r})\ell_2} \mathcal{M}''$ for another $\mathcal{M}''$. Now consider the reduction path $\rho = \mathcal{M} \xrightarrow{(\mathsf{p},\mathsf{q})\ell_1} \mathcal{M}'$. $(\mathsf{p},\mathsf{r})\ell_2$ is enabled at $\mathcal{M}$ so in a fair path it should eventually be executed, however no extension of ρ can contain such a transition as $\mathcal{M}'$ has no remaining transitions. Nevertheless, it turns out that there is a fair reduction path starting from every typable session (Lemma 6.14), and this will be enough to prove our desired liveness property.

We can now lift the typing relation to reduction paths, just like we did in Definition 5.19.

▶ **Definition 6.11** (Path Typing). *Path typing is a relation between session reduction paths and local type context reduction paths, defined coinductively by the following rules:*

(i) *The empty session reductoin path is typed with the empty context reduction path.*

(ii) *If* $\mathcal{M} \xrightarrow{\lambda_0} \rho$ *is typed by* $\Gamma \xrightarrow{\lambda_1} \rho'$ *where (*ρ *and* ρ' *are session and local type context reduction paths, respectively), then* $\lambda_0 = \lambda_1$ *and* ρ *is typed by* ρ'*.*

Similar to Lemma 5.20, we can show that if the head of the path is typable then so is the whole path.

▶ **Lemma 6.12.** *If* `typ_sess M gamma`*, then any session reduction path* `xs` *starting with* `M` *is typed by a local context reduction path* `ys` *starting with* `gamma`*.*

Proof. We can construct a local context reduction path that types the session path. The construction exactly like Lemma 5.20 but elements of the output stream are generated by Theorem 6.2 instead of Theorem 4.10. ◀

We also have that typing path preserves fairness.

▶ **Lemma 6.13.** *If session path* `xs` *is typed by the local context path* `ys`*, and* `xs` *is fair, then so is* `ys`*.*

The final lemma we need in order to prove liveness is that there exists a fair reduction path from every typable session.

▶ **Lemma 6.14** (Fair Path Existence). *If* `typ_sess M gamma`*, then there is a fair session reduction path* `xs` *starting from* `M`*.*

Proof. We can construct a fair path starting from `M` by repeatedly cycling through all participants, checking if there is a transition involving that participant, and executing that transition if there is. ◀

▶ Remark 6.15. The Rocq implementation of Lemma 6.14 computes a `CoFixpoint` corresponding to the fair path constructed above. As in Lemma 5.20, we use `constructive_indefinite_description` to turn existence statements in `Prop` to dependent pairs. We also assume the informative law of excluded middle (`excluded_middle_informative`) in order to carry out the "check if there is a transition" step in the algorithm above. When proving that the constructed path is fair, we sometimes rely on the LTL constructs we outlined in Section 5.2 reminiscent of the techniques employed in [4].

We can now prove that typed sessions are live.

▶ **Theorem 6.16** (Liveness by Typing). *For a session* `Mp`*, if* $\exists$ `gamma gamma` $\vdash_{\mathcal{M}}$ `Mp` *then* `live_sess Mp`*.*

Proof. We detail the proof for the send case of Definition 6.8, the case for the receive is similar. Suppose that `Mp` $\rightarrow^*$ `M` and `M` $\Rrightarrow$ `((p` $\leftarrow$ `p_send q ell e P') ||| M')`. Our goal is to show that there exists a `M''` such that `M` $\rightarrow^*$ `((p` $\leftarrow$ `P')|||M'')`. First, observe that by [R-Unfold] it suffices to show that `((p` $\leftarrow$ `p_send q ell e P') ||| M')` $\rightarrow^*$ `M''` for some `M''`. Also note that `gamma` $\vdash_{\mathcal{M}}$ `M` for some `gamma` by Theorem 6.2, therefore `gamma` $\vdash_{\mathcal{M}}$ `((p` $\leftarrow$ `p_send q ell e P') ||| M')` by Lemma 6.1.

Now let `xs` be a fair reduction path starting from `((p` $\leftarrow$`- p_send q ell e P') ||| M')`, which exists by Lemma 6.14. Let `ys` be the local context reduction path starting with `gamma` that types `xs`, which exists by Lemma 6.12. Now `ys` is fair by Lemma 6.13. Therefore by Theorem 5.24 `ys` is live, so a `lcomm p q ell'` transition eventually occurs in `ys` for some `ell'`. Therefore `ys` $=$ `gamma` $\rightarrow^*$ `gamma_0` $\xrightarrow{(\mathtt{p},\mathtt{q})\ell'}$ `gamma_1` $\rightarrow$ .. for some `gamma_0`, `gamma_1`. Now consider the session `M_0` typed by `gamma_0` in `xs`. We have `((p` $\leftarrow$ `p_send q ell e P') ||| M'')` $\rightarrow^*$ `M_0` by `M_0` being on `xs`. We also have that `M_0` $\xrightarrow{(\mathtt{p},\mathtt{q})\ell''}$ `M_1` for some ℓ'', `M_1` by Theorem 6.5. Now observe that `M_0` $\equiv$ `((p` $\leftarrow$ `p_send q ell e P') ||| M'')` for some `M''` as no transitions involving `p` have happened on the reduction path to `M_0`. Therefore $\ell = \ell''$, so `M_1` $\equiv$ `((p` $\leftarrow$ `P') ||| M'')` for some `M''`, as needed. ◀

7 Conclusion and Related Work

Liveness Properties. Examinations of liveness, also called *lock-freedom*, guarantees of multiparty session types abound in literature, e.g. [32, 24, 46, 37, 3]. Most of these papers use the definition liveness proposed by Padovani [31], which doesn't make the fairness assumptions that characterize the property [17] explicit. Contrastingly, van Glabbeek et. al. [43] examine several notions of fairness and the liveness properties induced by them, and devise a type system with flexible choices [7] that captures the strongest of these properties, the one induced by the *justness* [44] assumption. In their terminology, Definition 6.8 corresponds to liveness under strong fairness of transitions (ST), which is the weakest of the properties considered in that paper. They also show that their type system is complete i.e. every live process can be typed. We haven't presented any completeness results in this paper. Indeed, our type system is not complete for Definition 6.8, even if we restrict our attention to safe and race-free sessions. For example, the session described in [43, Example 9] is live but not typable by a context associated with a balanced global type in our system.

Fairness assumptions are also made explicit in recent work by Ciccone et. al [11, 12] which use generalized inference systems with coaxioms [1] to characterize *fair termination*, which is stronger than Definition 6.8, but enjoys good composition properties.

Mechanisation. Mechanisation of session types in proof assistants is a relatively new effort. Our formalisation is built on recent work by Ekici et. al. [15] which uses a coinductive representation of global and local types to prove subject reduction and progress. Their work uses a typing relation between global types and sessions while ours uses one between associated local type contexts and sessions. This necessiates the rewriting of subject reduction and progress proofs in addition to the operational correspondence, safety and liveness properties we have proved. Other recent results mechanised in Rocq include Ekici and Yoshida's [16] work on the completeness of asynchronous subtyping, and Tirore's work [39, 41, 40] on projections and subject reduction for π-calculus.

Castro-Perez et. al. [9] devise a multiparty session type system that dispenses with projections and local types by defining the typing relation directly on the LTS specifying the global protocol, and formalise the results in Agda. Ciccone's PhD thesis [10] presents an Agda formalisation of fair termination for binary session types. Binary session types were also implemented in Agda by Thiemann [38] and in Idris by Brady[6]. Several implementations of binary session types are also present for Haskell [25, 29, 36].

Implementations of session types that are more geared towards practical verification include the Actris framework [19, 22] which enriches the seperation logic of Iris [23] with binary session types to certify deadlock-freedom. In general, verification of liveness properties, with or without session types, in concurrent seperation logic is an active research area that has produced tools such as TaDa [14], FOS [26] and LiLo [27] in the past few years. Further verification tools employing multiparty session types are Jacobs's Multiparty GV [22] based on the functional language of Wadler's GV [45], and Castro-Perez et. al's Zooid [8], which supports the extraction of certifiably safe and live protocols.

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