Automated Modular Verification for Race-Free Channels with Implicit and Explicit Synchronization

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Abstract

Ensuring the correctness of software for communication-centric programs is important but challenging. Previous approaches, based on session types, have been intensively investigated over the past decade. They provide a concise way to express protocol specifications and a lightweight approach for checking their implementation. Current solutions are based on only implicit synchronization, and are based on the less precise types rather than logical formulae. In this paper, we propose a more expressive session logic to capture multi-party protocols. By using two kinds of ordering constraints, namely “happens-before” \( \prec_{\text{HB}} \) and “communicates-before” \( \prec_{\text{CB}} \), we show how to ensure from first principle race-freedom over common channels. Our approach refines each specification with both assumptions and proof obligations to ensure compliance to some global protocol. Each specification is then projected for each party and then each channel, to allow cooperative proving through localized automated verification. Our primary goal in automated verification is to ensure race-freedom and communication-safety, but the approach is extensible for deadlock-freedom as well. We shall also describe how modular protocols can be captured and handled by our approach.

1 Introduction

The notable performance and scalability of our computing platform can be attributed to the virtue of distributed computation. In turn, distributed computation has seen a rise in the level of concurrency with the adoption of the message passing model, where concurrent programs communicate with each other via communication channels. Given the somewhat complex interaction schemes used, ensuring the safety and correctness of such system is challenging. Some research progress towards specification-based development for the message passing paradigm have mostly revolved around the session-based concurrency. In these approaches, the communication patterns are abstracted as session types for variants of \( \pi \)-calculus [20, 35] which are either statically checked against given specifications or are being used to guide the generation of deadlock-free code [13].

While providing a foundational approach towards communication correctness, most of these works make the assumption that the underlying system communicates exclusively via message passing using unlimited channels. That is, extra channels are created when necessary to resolve any ambiguity in the communication protocol. This assumption is broken in current systems which may rely on more tightly synchronized concurrency. Concurrency is often managed by using a combination of synchronization mechanisms which are designed to work together for better performance over shared channels. For instance, the Android systems recognize the benefits of using message passing, but are not hesitant in combining this style with other synchronization mechanisms such as CountDownLatch. Moreover, while sequential and disjunctive combinations of communication channels are quite well understood, the same cannot be said for concurrent compositions of transmissions over shared channels.

In summary, most of the current approaches which guarantee communication safety, conveniently limit their results to implicit synchronization of protocols. Should race-related conflict arises, extra transmission channels are arbitrary added to ensure unambiguous communication. In this paper, we consider how to abstractly extend communication protocols to those which require explicit synchronization. Our thesis is that specification in concurrency should be abstract but with sufficient detail so that implementation can be safely written to meet the intention of protocol designers. We aim for an expressive logic for session-based concurrency; and a set of tools for automated reasoning of our enhanced session logic. We consider how to ensure well-formedness of session logic, and how to build summaries to ensure race-freedom and type safety amongst shared channels.

As an example, session types approaches only impose sequencing on the actions of the same participant, as illustrated here: \((A \rightarrow C) \land (B \rightarrow C)\). This example involves three communicating parties with \(C\) receiving data from \(A\) and \(B\), respectively. For simplicity, we assume asynchronous communication with non-blocking send and blocking receive.
We further assume that messages are received in a FIFO order in each of these communication channels. In terms of communications, the two sends (by A and B) are unordered, while the two receives by C are strictly sequentialized (by C itself) with the receipt of message from A expected to occur before the message from B. Such implicit synchronization may only work if two distinct channels are being used to serve the transmissions of \text{A} \rightarrow \text{C} and that of \text{B} \rightarrow \text{C}, respectively, since the two sends by A and B would then arrive unambiguously, irrespective of their arrival order over the two distinct channels, say c_1 and c_2, as shown next: \((A \rightarrow \text{C} : c(t_1)) \land (B \rightarrow \text{C} : c(t_2))\). However, if we had to use a common channel, say c for mailbox of party C, we would have the following specification.

\[(A \rightarrow \text{C} : c(t_1)) \land (B \rightarrow \text{C} : c(t_2))\]

This is perfectly legitimate depending on the programming models used for implementing the given protocol. For example, given the following two code snippets, the second implementation (b) of the above protocol is safe, while the first one (a) contains a race potentially leading to an unsafe communication:

(a)

\[
\begin{array}{ccc}
\text{A} & \text{B} & \text{C} \\
\text{send}(c, 'Lorelei') & \text{send}(c, 68) & \text{book = recv}(c) \\
\text{price = recv}(c) & & \\
\end{array}
\]

(b)

\[
\begin{array}{ccc}
\text{A} & \text{B} & \text{C} \\
\text{send}(c, 'Lorelei') & \text{wait}(w) & \text{notifyAll}(w) \\
\text{send}(c, 68) & \text{book = recv}(c) & \text{price = recv}(c) \\
\end{array}
\]

We also provide unique labels \(i_1\) and \(i_2\) for each transmission, and its corresponding message types \(t_1\) and \(t_2\). We assume each party never sends message to itself. This allows us to use \(\text{R}(i)\) to unambiguously refer to a send or receive event by party \(R\) at a transmission labelled \(i\). We now have a situation where two senders (by A and B) must be explicitly ordered to have their messages arrive in strict sequential order. (The two receives by the same party C are already implicitly ordered.) It may even become a communication safety issue if \(t_1 \neq t_2\) and should the two messages arrive in the wrong order. We refer to this problem as a channel race where messages could be sent to unintended destinations. In this case, we have to arrange for the two sends (initiated by A and B) to be strictly ordered. For this to be supported, we propose an explicit synchronization mechanism that would force the second send by B to occur after the first send by A has been initiated. While explicit synchronization can be handled by a number of mechanisms, such as \text{notifyAll} − \text{wait} or \text{CountDownLatch}, our specification logic shall abstractly capture this requirement by an \text{ordering}, namely \(A^{(i_1)} \preceq_{\text{IB}} B^{(i_2)}\), which requires that event \(A^{(i_1)}\) to happen before event \(B^{(i_2)}\).

Such explicit orderings should be minimised, where possible, but they are an essential component of concurrency control for the message passing paradigm. Our contributions are:

- We design Mercurius, an expressive session logic that is both precise and concise for modeling multi-party protocols. Our approach utilizes two fundamental ordering constraints.
- We ensure race-freedom in communications over common channels with both implicit and explicit synchronization. Prior works, based on session types, relied on only implicit synchronization.
- We provide a specification refinement that explicitly introduces both assumptions and proof obligations. These may be either local or global events and proofs, but can be used together to ensure adherence to the global protocol, communication safety and race-freedom.
- Each refined specification may be firstly projected for each party and then for each channel, with a set of shared global assumptions. There are three novelties. First, global proof obligations may be projected to support cooperative proving, where the proof by one party can be relied as assumption by the other concurrent parties. Second, we use event guards to ensure that channel specifications are verified in correct sequence. Lastly, multi-endpoints for channels are supported.
- We show how automated verification is applied on a per-party basis with global assumptions and proof obligations. Assumptions are released as soon as possible. Proof obligations can be delayed and locally proven at the appropriate time.
- We propose an approach to express global protocols in a modular fashion, where "plug-and-play" protocols are only designed and refined once, and re-used multiple times in different contexts. This greatly facilitates specification re-use and forms the basis for supporting both inductive protocol specification and their recursive implementation.

Limitation: For simplicity, we currently restrict our protocol to well-formed disjunction where a common sender and receiver is always and exclusively present in each disjunction to express communications across its branches. This restriction provides a simplification where every \(\preceq_{\text{IB}}\) ordering discovered is a must-ordering, since transitivity holds globally whenever it holds in one of the branches. Lifting this restriction will require orthogonal mechanisms to be supported, such as partial orderings and synchronization across multiple parties in conditional branches.

2 Overview

To capture a broad range of communication patterns, we propose an expressive logic for specification of global protocols in Fig. 1a. This protocol uses \(G_1 \ast G_2\) for the concurrency
Automated Modular Verification for Race-Free Channels with Implicit and Explicit Synchronization

of \( G_1 \) and \( G_2 \), and \( G_1 \lor G_2 \) for disjunctive choice between either \( G_1 \) and \( G_2 \), and finally \( G_1 \land G_2 \) on the implicit sequentialization of \( G_1 \) before \( G_2 \) for either the same party or same channel communication. Moreover, we express the message type of a transmission \( S \xrightarrow{t} R : c(v \cdot \Delta) \) by \( v \cdot \Delta \) which expects a message \( v \) on some resource expressed in logical form \( \Delta \) (defined in Fig. 1b) transmitted over channel \( c \). As an example, \( v \cdot 11(v, n) \) would capture the transmission of half fractional permission\(^1\) of a linked list with \( n \) elements \( 11(v, n) \) rooted at \( v \). Apart from messages in logical form, our session logic captures logical disjunction and can enforce explicit synchronization to ensure communication safety.

We provide a specification refinement that adds unique transmission labels \( i \), guards (of the form \( \equiv(\Psi) \)) and assumptions (of the form \( \equiv(\Psi) \)), as more precise denotation of each global protocol. Guards capture assertions (or proof obligations), such as explicit synchronization, for its global protocol to ensure race-free communications; while assumptions support flow of information across multiple parties to facilitate local verification. To sketch the main purpose of using these guards and assumptions, we briefly introduce some elements of \( \Psi \) later in this section.

These new specification constructs are added automatically by our refinement procedure to help ensure race-freedom in shared channels. Our refinement could also add channel information based on some expected communication patterns. For example, if mailbox communication style is preferred, each communication always directs its message to the mailbox of its receiver. If bi-directional channels of communication are preferred, a distinct channel is provided for every pair of parties that are in communication. To support diversity, we assume channels are captured explicitly in our protocol specifications.

Our approach thus takes in a protocol specified in an expressive logic, refine it, then project it on a per party and per channel basis. Once these are done, we can proceed to apply automated verification on the entire program code using the derived specs.

The automated verification is conducted against a general and expressive specification language based on Separation Logic [15, 61], as shown in Fig. 1b, which supports the use of (inductive) predicates, the specification of separation and numerical properties, as well as channel predicates.

2.1 Specification Refinement

Let us use a well-known example to illustrate the key ideas of our approach. Consider the multi-party protocol from [35] where we use a distinct channel as mailbox of each of the three parties, namely seller \( S \), first buyer \( B_1 \) and the second collaborative buyer \( B_2 \).

\[
\begin{align*}
B_1 & \rightarrow S : s(Order) ; (S \rightarrow B_1 : b_1.Price) \lor (S \rightarrow B_2 : b_2.Price)) ; B_1 \rightarrow B_2 : b_1.Amt) ; \\
(B_2 \rightarrow S : s(No) \lor (B_2 \rightarrow S : s(Yes)) ; B_2 \rightarrow S : s(Addr))
\end{align*}
\]

Note that \( T \) is a short-hand for \( v \cdot T(v) \) where \( T \) is a type or predicate for some resource, such as \( \equiv(Order, Price, Amt \) or \( Addr) \).

For the above example, our refinement would add information to signify the timely completion of each communication completed using \( \equiv(S \xrightarrow{t} R) \). This event itself captures two locally observable events and a globally observable ordering, as follows:

\[
\equiv(S \xrightarrow{t} R) \Rightarrow \equiv(S^{(i)} \land \equiv(R^{(i)}) \land \equiv(S^{(s)}<_{CB}R^{(s)})
\]

which indicates the occurrence of the send and receive events and the expected \textit{communicates-before} ordering\(^2\), denoted by \( <_{CB} \), for its successfully completed communication. The send and receive events are only observable locally. For example, the completion of event \( \equiv(P_1^{(i)}) \) is only observable in party \( P_1 \), but not by other distinct parties, such as \( P_2 \). Its occurrence in the global protocol is thus location sensitive and can be used to capture relative orderings within the same party. However, the ordering \( \equiv(S^{(i)}<_{CB}R^{(i)}) \) is globally observable since it relates two parties and can rely on our assurance that all send/receive events are uniquely labelled.

We also generate immediate \( <_{IB} \) orderings between each pair of adjacent events of the same protocol in the global protocol, as follows:

\[
\equiv(P_1^{(i)}) \Rightarrow \equiv(P_2^{(i)}) \Rightarrow \equiv(P_3^{(i)}) \Rightarrow \equiv(P_4^{(i)}) \Rightarrow \equiv(P_5^{(i)}) \Rightarrow \equiv(P_6^{(i)}) \Rightarrow \equiv(P_7^{(i)}) \Rightarrow \equiv(P_8^{(i)}) \Rightarrow \equiv(P_9^{(i)})
\]

When common channels are involved, we also use proof obligations of the form \( \equiv(\leq_{\text{IB}}) \) to denote two transmissions, labelled \( i_1 \) and \( i_2 \), that are expected to be ordered to ensure race freedom. Assuming each transmission \( i \) be denoted by \( S_1 \xrightarrow{t_i} R_1 : c(v \cdot \Delta_i) \) with \( c \) as common channel, these proof obligations can be further reduced, as follows:

\[
\equiv(S_1^{(i)} \equiv_{\text{IB}} S_2^{(i)}) \land \equiv(R_1^{(s)} \equiv_{\text{IB}} R_2^{(s)}))
\]

The \( <_{\text{IB}} \) is overloaded for both transmissions e.g. \( i_1 \equiv_{\text{IB}} i_2 \) and events e.g. \( P_1^{(i_1)} \equiv_{\text{IB}} P_2^{(i_2)} \). Our global specification \( G \) is now refined to \( G @ \text{Refine} \) as shown below.

\[
G @ \text{Refine} \equiv \begin{cases}
B_1 \rightarrow S : s(Order) ; (B_1 \rightarrow S) ; \\
((S \rightarrow B_1 : b_1.Price ; (S \rightarrow B_1 ; c(S^{(1)}<_{\text{IB}} S^{(2)}))) ; (B_1 \rightarrow B_2 : b_1.Amt) ; (B_1 \rightarrow B_2 ; c(B_1^{(1)}<_{\text{IB}} B_2^{(1)}) ; c(T^{(1)}<_{\text{IB}} T^{(2)})) ; (B_2 \rightarrow S : s(No) ; (B_2 \rightarrow S ; c(B_2^{(4)}<_{\text{IB}} B_2^{(5)}) ; c(S^{(2)}<_{\text{IB}} S^{(3)})) ; (B_2 \rightarrow S ; c(S^{(s)}<_{\text{IB}} S^{(s)})) ; \equiv(S^{(s)}<_{\text{IB}} S^{(s)})) ; \equiv(1<_{\text{IB}} 2)) \lor \\
B_2 \rightarrow S : s(Yes) ; (B_2 \rightarrow S ; c(B_2^{(4)}<_{\text{IB}} B_2^{(5)}) ; c(S^{(2)}<_{\text{IB}} S^{(3)})) ; (B_2 \rightarrow S ; c(S^{(s)}<_{\text{IB}} S^{(s)})) ; \equiv(S^{(s)}<_{\text{IB}} S^{(s)})) \lor \equiv(1<_{\text{IB}} 2)) \lor \\
B_2 \rightarrow S : s(Addr) ; (B_2 \rightarrow S ; c(B_2^{(4)}<_{\text{IB}} B_2^{(5)}) ; c(S^{(s)}<_{\text{IB}} S^{(s)})) \lor \equiv(1<_{\text{IB}} 2)) \lor \\
\end{cases}
\]

\(^1\)Fractional permissions are important for supporting concurrent programming. Its use allows read access when a resource is shared by multiple parties, but also write access when a resource later becomes exclusively owned.

\(^2\)Under race-free assumption, \( <_{\text{IB}} \) ordering is the same as \( <_{\text{IB}} \)-ordering.
2.2 Specification Projection

Once we have a specification that has been annotated with assumptions and obligations, we can proceed to project them for use in code verification. The projection can be done in two phases. Firstly, on a per-party basis, and secondly on a per-channel basis. A per-channel within each party specification allows us to observe all communication and proof obligation activities within each channel. We propose to undertake per-channel projections on specifications for two reasons. Firstly, it supports backward compatibility, as an earlier two-party session logic verification system, [18], was designed based on per-channel specifications. Secondly, it is more natural for communication primitives, such as send/receive, to be denoted by their channel’s specifications. Lastly, we also project assumptions on global orderings, so that they may be utilized by each party for their respective local verification.

Let us first discuss the specification that would be projected for each party. For the running example, we would project the following local specifications for the three parties.

\[ G\#P_1 = \{s\text{Order}; @B(1)\}; b_1\text{Price}; @B(2); b_2\text{Am'}t; @B(4); @3\text{-eq}4\} \]

\[ G\#P_2 = \{b_2\text{Price}; @B(1); b_2\text{Am'}t; @B(4); @3\text{-eq}4\} \]

\[ G\#S = \{s\text{Order}; @S(1); (b_1\text{Price}; @S(2)) \od (b_2\text{Price}; @S(3)) ; @3\text{-eq}4\} \]

Each of the local specification contains send/receive proof obligations, such as s\text{Order} and s\text{Order}, which capture expected channel and message type (or property). It also contains assumptions on the local events, such as @B(1) and @S(1) that are immediately released once the corresponding send/receive obligations have been locally verified. We may also have global proof obligation, such as @3\text{-eq}4, that is meant to ensure that the two transmissions, labelled as 3 and 4, are race-free since they share a common channel. Our projection of each global proof obligation to multiple parties supports the concept of cooperative proving, whereby the proving effort by one party is utilized as an assumption by the other parties. Thus, @3\text{-eq}4 is translated to the following cooperative proving and assumptions:

\[ \Theta(\{s\text{Order}; @B(1); b_1\text{Price}; @B(2); b_2\text{Am'}t; @B(4)\}) = (\Theta(B(1)) <_{GB} \Theta(S(1)) \land \Theta(B(2)) <_{GB} \Theta(B(1)) \land \Theta(B(4)) <_{GB} \Theta(B(2))) \]

\[ \Theta(\{s\text{Order}; @B(1); b_1\text{Price}; @B(2); b_2\text{Am'}t; @B(4)\}) = (\Theta(B(1)) <_{GS} \Theta(S(1)) \land \Theta(B(2)) <_{GS} \Theta(S(1)) \land \Theta(B(4)) <_{GS} \Theta(B(2))) \]

\[ \Theta(\{s\text{Order}; @B(1); b_1\text{Price}; @B(2); b_2\text{Am'}t; @B(4)\}) = (\Theta(B(1)) <_{GB} \Theta(B(2)) \land \Theta(B(3)) <_{GB} \Theta(B(1)) \land \Theta(B(4)) <_{GB} \Theta(B(2))) \]

\[ \Theta(\{s\text{Order}; @B(1); b_1\text{Price}; @B(2); b_2\text{Am'}t; @B(4)\}) = (\Theta(S(1)) <_{GS} \Theta(S(2)) \land \Theta(S(1)) <_{GS} \Theta(S(3)) \land \Theta(S(2)) <_{GS} \Theta(S(3))) \]

The other global orderings from each transmission, such as communicates-before ordering like @B(1) <_{GB} S(1), and happens-before ordering between adjacent events within the same party, like @S(1) <_{GS} S(2), are placed in a shared space, as say G\#All, that is visible to all parties. As these global ordering information are modelled for unique events and are also immutable, we propose to release all orderings in a single step at the beginning of the protocol for simplicity. Our proposal to release all global ordering together as a single assumption is done for simplicity. Though future orderings are not required to prove current global obligations, they never cause any inconsistency to the current state, but merely add some new orderings that are available in advance.

\[ G\#P_1 = \{s\text{Order}; @B(1); b_1\text{Price}; @B(2); b_2\text{Am'}t; @B(4)\} \land \Theta(\{s\text{Order}; @B(1); b_1\text{Price}; @B(2); b_2\text{Am'}t; @B(4)\}) \]

\[ G\#P_2 = \{b_2\text{Price}; @B(1); b_2\text{Am'}t; @B(4)\} \land \Theta(\{b_2\text{Price}; @B(1); b_2\text{Am'}t; @B(4)\}) \]

\[ G\#S = \{s\text{Order}; @S(1); (b_1\text{Price}; @S(2)) \od (b_2\text{Price}; @S(3)) ; @3\text{-eq}4\} \land \Theta(\{s\text{Order}; @S(1); (b_1\text{Price}; @S(2)) \od (b_2\text{Price}; @S(3)) ; @3\text{-eq}4\}) \]

Given global specification G over n parties, P_1 \cdots P_n, our projection would transform G into a per-party specification, as follows:

\[ G \Rightarrow G\#P_1 \cdots \Rightarrow G\#P_n \Rightarrow G\#All \]

where G\#P_1 \cdots \Rightarrow G\#P_n denote the per-party specifications and G\#All denotes the global orderings that are shared by all parties. Once we have a per-party specification, say G\#P_1 over channels c_1 \cdots c_n, we can further project each of these specifications into a spatial conjunction of several per-channel specifications, namely G\#P\#c_1 \cdots \Rightarrow G\#P\#c_m, as follows:

\[ G\#P_1 \Rightarrow G\#P\#c_1 \cdots \Rightarrow G\#P\#c_m \]

The details for channel projections are given later in Sec 4.
2.3 Automated Local Verification

Once we have the refined specification that had been suitably projected on a per party and per channel basis, we can proceed to use these specifications for automated local verification. The shared global orderings can be immediately added to the initial program state. The per-channel specification are released during the verification on each of the respective program codes, namely $\text{Code}_1 \parallel \text{Code}_2 \parallel \text{Code}_3$, for the different parties. For our running example, the initial and final program states are:

\[
\{\text{Common}(\#\text{All}) \ast \text{Party}(S, \#S) \ast \text{Party}(B_1, \#B_1) \ast \text{Party}(B_2, \#B_2)\} \\
\{\text{Party}(S, \#S) \ast \text{Party}(B_1, \#B_1) \ast \text{Party}(B_2, \#B_2) \wedge \#\text{All}\}
\]

The $\text{emp}$ in the final state of each party can help to confirm the completion of all transmissions in the global protocol. The abstract predicate $\text{Party}(S, \#S)$ associates each party with its corresponding specification. Our approach supports both events that are either immutable or mutable. Each immutable event (seen earlier) is labelled uniquely, while mutable events can be updated flow-sensitively. To signify each executing party $P$, we use a mutable event $\text{Peer}(P)$ that is updated when $P$ is executing. Such mutable events are added to as ghost specifications, as shown below:

\[
\{\text{Common}(\#\text{All}) \ast \text{Party}(S, \#S) \ast \text{Party}(B_1, \#B_1) \ast \text{Party}(B_2, \#B_2)\} \\
\{\text{Party}(S, \#S) \ast \text{Party}(B_1, \#B_1) \ast \text{Party}(B_2, \#B_2) \wedge \#\text{All}\}
\]

\[
\{\text{Common}(\#\text{All}) \ast \text{Party}(S, \#S) \ast \text{Party}(B_1, \#B_1) \ast \text{Party}(B_2, \#B_2) \} \\
\{\text{Party}(S, \#S) \wedge \#\text{All}\} \oplus (\text{Peer}(S)) ; \text{Code}_S \{\text{Party}(S, \#S)\}
\]

\[
\{\text{Party}(B_1, \#B_1) \wedge \#\text{All}\} \oplus (\text{Peer}(B_1)) ; \text{Code}_B \{\text{Party}(B_1, \#B_1)\}
\]

\[
\{\text{Party}(B_2, \#B_2) \wedge \#\text{All}\} \oplus (\text{Peer}(B_2)) ; \text{Code}_B \{\text{Party}(B_2, \#B_2)\}
\]

\{
\text{Party}(S, \#S) \ast \text{Party}(B_1, \#B_1) \ast \text{Party}(B_2, \#B_2) \}

Note that $\text{Common}(\#\text{All})$ denotes pure global information that can be duplicated and propagated.

3 Global Protocols

We now formalize our proposal into a logical system called Mercurius, whose syntax is depicted in Fig. 1a. We first list down the elements of the protocol and their properties before studying the properties of the whole protocol.

**Communication model.** To support a wide range of communication interfaces, the current session logic is designed for a permissive communication model, where:

- The transfer of a message dissolves asynchronously, that is to say that sending is non-blocking while receiving is blocking.
- The communication interface of choice manipulates **FIFO channels** that are shareable, in the sense that each channel can serve two or more participants.
- For simplicity, the communication assumes unbounded buffers. However, extension to bounded buffer is possible due to our projection to a per-channel specification.

**Transmission.** As depicted in Fig. 1a, a transmission $S \rightarrow c (v \cdot \Delta)$ involves a sender $S$ and a receiver $R$ transmitting a message $v$ expressed in logical form $\Delta$ over a buffered channel $c$. This transmission is uniquely identified by a label $i$. In the subsequent we often use only the unique label $i$ to refer to a particular transmission. To access the components of a transmission we define the following auxiliary functions: $\text{send}(i) \triangleq S(i)$, $\text{recv}(i) \triangleq R(i)$, $\text{chan}(i) \triangleq c$ and $\text{msg}(i) \triangleq v \cdot \Delta$. A transmission is irreflexive, since it would make no sense for the sending and the receiving to be performed by the same peer. We define a function $\text{TR}(G)$ which decomposes a given protocol to collect a set of all its constituent transmissions, and a function $\text{TR}_{\text{fst}}(G)$ to return the set of all possible first transmissions:

\[
\text{TR}(S \rightarrow c (v \cdot \Delta)) \triangleq \{S \rightarrow c (v \cdot \Delta)\}
\]

\[
\text{TR}(G_1 \cup G_2) \triangleq \text{TR}(G_1) \cup \text{TR}(G_2)
\]

\[
\text{TR}(\oplus(i)) \triangleq \emptyset
\]

\[
\text{TR}(G_1 \vee G_2) \triangleq \text{TR}(G_1) \cup \text{TR}(G_2)
\]

\[
\text{TR}_{\text{fst}}(G_1 \cup G_2) \triangleq \text{TR}_{\text{fst}}(G_1) \cup \text{TR}_{\text{fst}}(G_2)
\]

\[
\text{TR}_{\text{fst}}(G_1 \vee G_2) \triangleq \text{TR}_{\text{fst}}(G_1) \cup \text{TR}_{\text{fst}}(G_2)
\]

**Event.** An event $E$ is a pair $P(i)$ where $P \in \text{Role}$ is the sending or receiving party of a transmission identified by $i \in \text{Nat}$. Given the uniqueness of the identifier $i$, an event uniquely identifies a send or receive operation. We denote by $E, \text{peer}$ and $E, \text{id}$ the elements of an event, e.g. $P(0).\text{peer} \triangleq P$ and $P(0).\text{id} \triangleq i$. The following function collects the set of all the events within a protocol:

\[
\text{EV}(G_1 \cup G_2) \triangleq \text{EV}(G_1) \cup \text{EV}(G_2)
\]

\[
\text{EV}(\oplus(i)) \triangleq \text{EV}(\oplus(i))
\]

\[
\text{EV}(G_1 \vee G_2) \triangleq \text{EV}(G_1) \cup \text{EV}(G_2)
\]

\[
\text{EV}(P(i)) \triangleq \{P(i)\}
\]

\[
\text{EV}(S \rightarrow c (v \cdot \Delta)) \triangleq \{S(i), R(i)\}
\]

\[
\text{EV}(P(1) \rightarrow c (P(2))) \triangleq \{P(1), P(2)\}
\]

\[
\text{EV}(P(1) \rightarrow c (P(2))) \triangleq \{P(1), P(2)\}
\]

The messages of two arbitrary but distinct transmissions, say $i_1$ and $i_2$, are said to be disjoint, denoted $\text{msg}(i_1) \neq \text{msg}(i_2)$ if $\text{UNSAT}(\Delta_1 \wedge \{v_1 \wedge v_2\} \Delta_2)$, where $\text{msg}(i_1) = v_1 \cdot \Delta_1$ and $\text{msg}(i_2) = v_2 \cdot \Delta_2$.

We abuse the set membership symbol, $\epsilon$, to denote the followings:
when edges represent the sequence relation between these nodes, $G$ and sequence form semigroups, $G(\cdot)$ is a directed acyclic graph (DAG).

We capture global protocol in terms of $emp$, Andreea Costea, Chin Wei-Ngan, Florin Craciun, and Shengchao Qin

$∀emp \equiv emp \equiv emp \equiv G$

Sequential composition is not commutative, unless it satisfies certain disjointness properties:

$G_1 : G_2 \equiv G_1 : G_2$

when $∀c_1 \in G_1, c_2 \in G_2 \Rightarrow c_1 \neq c_2$ and $∀P_1 \in G_1, P_2 \in G_2 \Rightarrow P_1 \neq P_2$

Sub-protocol. A protocol $G'$ is said to be a sub-protocol of $G$ when $G'$ is a decomposition of $G$, where the set of all possible decompositions of $G$ is recursively defined on the structure of $G$ as follows:

$\text{sub}(S \rightarrow R : c(\cdot \cdot \cdot)) \equiv \text{sub}(S \rightarrow R : c(\cdot \cdot \cdot))$

$\text{sub}(\oplus(\cdot \cdot \cdot)) \equiv \emptyset$

$\text{sub}(\oplus(\cdot \cdot \cdot)) \equiv \emptyset$

$\text{sub}(G_1; G_2) \equiv \{G_1; G_2\} \cup \text{sub}(G_1) \cup \text{sub}(G_2)$

$\text{sub}(G_1 \lor G_2) \equiv \{G_1 \lor G_2\} \cup \text{sub}(G_1) \cup \text{sub}(G_2)$

$\text{sub}(G_1 \lor G_2) \equiv \{G_1 \lor G_2\} \cup \text{sub}(G_1) \cup \text{sub}(G_2)$

Graph ordering. We capture global protocol in terms of a directed acyclic graph $G(\cdot) = (V, O)$, where the set of vertices is the set of all the transmissions in $G$, $V = TR(G)$, and the edges represent the sequence relation between these nodes, $P = \text{seq}(G)$, where $\text{seq}$ is defined as follows:

$P \in G \Rightarrow \exists i \in G : c \in i$

$P \in G \Rightarrow \exists i \in G : c \in i$

$P \in G \Rightarrow \exists i \in G : c \in i$

$P \in G \Rightarrow \exists i \in G : c \in i$

Correspondingly, $\notin$ is used to denote the negation of the above.

Transmissions are organized into a global protocol using a combination of parallel composition $G_1 \lor G_2$, disjunction $G_1 \lor G_2$ and sequential composition $G_1 \land G_2$. The parallel composition of global protocols forms a commutative monoid $(G, \cdot, emp)$ with $emp$ as identity element, while disjunction and sequence form semigroups, $(G, \lor)$ and $(G, ;)$, with the former also satisfying commutativity. $emp$ acts the left identity element for sequential composition:

$G_1 \land G_2 \equiv G_2 \land G_1$

$G \lor emp \equiv emp \equiv G$

Two transmissions, $i_1$ and $i_2$, are sequenced in a protocol $G$, denoted $i_1 < i_2$, if there is a path in $G(\cdot)$ from $i_1$ to $i_2$. Two transmissions, $i_1$ and $i_2$, are adjacent in $G$ if they share the same channel $c$, they are sequenced in $G(\cdot)$, and there are no other transmission on $c$ in between $i_1$ and $i_2$. And lastly, two transmissions are linked if they are sequenced and they share the same channel. These relations are formally described in Fig. 2. Since these relations are defined as edges of a directed acyclic group it is straightforward to show that they are irreflexive and antisymmetric. The transitivity of sequenced also follows directly from the reachability relation of DAGs which is the transitive closure of the edges in $G(\cdot)$. A simple case analysis on the definition of linked transmissions shows that the linked relation is also transitive.

3.1 Well-Formedness

Concurrency. The $\cdot$ operator offers support for arbitrary-ordered (concurrent) transmissions, where the completion order is not important for the final outcome.

Definition 1 (Well-Formed Concurrency). A protocol specification, $G_1 \lor G_2$, is said to be well-formed with respect to $\cdot$ if and only if $∀c \in G_1 \Rightarrow c \not\in G_2$, and vice versa.

This restriction is to avoid non-determinism from concurrent communications over the same channel.

Choice. The $\lor$ operator is essential for the expressiveness of Mercurius, but its usage must be carefully controlled:

Definition 2 (Well-Formed Choice). A disjunctive protocol specification, $G_1 \lor G_2$, is said to be well-formed with respect to $\lor$ if and only if all of the following conditions hold, where $T_1$ and $T_2$ account for all first transmissions of $G_1$ and $G_2$, respectively, $T_1 = TR^{\text{fast}}(G_1)$ and $T_2 = TR^{\text{fast}}(G_2)$:

(a) (same first channel) $∀i_1, i_2 \in T_1 \lor T_2 \Rightarrow \text{chan}(i_1) = \text{chan}(i_2)$;

(b) (same first sender) $∀i_1, i_2 \in T_1 \lor T_2 \Rightarrow \text{send}(i_1).\text{peer} = \text{send}(i_2).\text{peer}$;

Andreea Costea, Chin Wei-Ngan, Florin Craciun, and Shengchao Qin

Figure 2. Transmission sequencing with respect to a given protocol $G$

$\text{seq}(G_1; G_2) = \{ \{i_1, i_2\} | i_1 \in TR(G_1) \land i_2 \in TR(G_2) \} \cup \text{seq}(G_1) \cup \text{seq}(G_2)$.

$\text{seq}(\oplus(\cdot \cdot \cdot)) = \emptyset$. $\text{seq}(\oplus(\cdot \cdot \cdot)) = \emptyset$. $\text{seq}(S \rightarrow R : c(\cdot \cdot \cdot)) = \emptyset$.

$\text{seq}(G_1 \lor G_2) = \text{seq}(G_1) \cup \text{seq}(G_2)$.

$\text{seq}(G_1 \lor G_2) = \text{seq}(G_1) \cup \text{seq}(G_2)$. 
Automated Modular Verification for
Race-Free Channels with Implicit and Explicit Synchronization

(c) (same first receiver) \( \forall i_1, i_2 \in T_1 \cup T_2 \Rightarrow \text{recv}(i_1).\text{peer} = \text{recv}(i_2).\text{peer} \);
(d) (mutually exclusive "first" messages) \( \forall i_1, i_2 \in T_1 \cup T_2 \Rightarrow \text{msg}(i_1) \# \text{msg}(i_2) \); \( i_1 = i_2 \);
(e) (same peers) \( \forall i_1, i_2 \in G_1 \cup G_2 \Rightarrow \{ \text{send}(i_1).\text{peer}, \text{recv}(i_1).\text{peer} \} = \{ \text{send}(i_2).\text{peer}, \text{recv}(i_2).\text{peer} \} \);
(f) (recursive well-formedness) \( G_1 \) and \( G_2 \) are well-formed with respect to \( \forall \).

**Definition 3** (Well-Formed Protocol). A protocol \( G \) is said to be well-formed, if and only if \( G \) contains only well-formed concurrent sub-protocols, and well-formed choices.

To ensure the correctness of our approach, Mercurius disregards as unsound any usage of \( * \) or \( \forall \) which is not well-formed.

### 3.2 Protocol Safety with Refinement

As highlighted in the overview of this paper, even simple protocols which only involve two transmissions can easily lead to flimsy communication, where either the receivers, \((A \rightarrow C ; B \rightarrow C)\), or the senders, \((A \rightarrow B ; A \rightarrow C)\), race for the same channel.

To avoid such race conditions it is essential to study the linearity of the communication. A communication is said to be linear when all the shared channels are used in a linear fashion, or in other words when a send and its corresponding receive are temporally ordered. Since the communication implementation is guided by the communication protocol, it is therefore essential for the protocol to satisfy this safety principle as well. In the subsequent we proceed in:

- defining a minimal set of causality relations relevant in the study of linearity and build an ordering system to reason about these relations.
- defining race-freedom w.r.t. these causality relations in the context of asynchronous communication.
- transform any given protocol into a race-free protocol via a refinement phase which adds assumptions about the causal relation between ordered events, and guards to enforce race-free communication.

**Ordering Constraint System.** Given a set of events \( E \) and a relation \( R \subseteq E \times E \), we denote by \( E_1 <_R E_2 \) the fact that \((E_1, E_2) \in R \). We next distinguish between two kinds of relations the so called "happens-before" relations to reflect the temporal ordering between events, and "communicates-before" relations to relate communicating peers.

**Definition 4** (Communicates-before). A communicates-before relation \( CB \) is defined as:

\[ \{ (E_1, E_2) \mid \exists i \cdot E_1.i.d = E_2.i.d = i \text{ and } E_1 = \text{send}(i) \text{ and } E_2 = \text{recv}(i) \} \]

The \( CB \) relation is not transitive, since its purpose is to simply assimilate the transitivity of HB by relating events which involve different peers. The \( CB \) is however irreflexive since we do not allow for a message to be sent and received by the same party, and antisymmetric since each event is unique (described by the \( \text{Role} \times \text{Nat} \) pair), that is, it only occurs within a single transmission.

**Definition 5** (Happens-before). Given a global protocol \( G \), two events \( E_1 \in EV(G) \) and \( E_2 \in EV(G) \) are said to be in a happens-before relation in \( G \) if \((E_1.i.d, E_2.i.d) \in \text{seq}(G)\) or it can be derived via a closure using propagation rules [HB-HB] and [CB-HB]. This is denoted by \( E_1 \triangleleft_{\text{HB}} E_2 \).

Similar to the happened-before a la Lamport [45], the HB relation is transitive, irreflexive and asymmetric, therefore HB is a strict partial order over events. However, one novel aspect of our approach is the use of [CB-HB] propagation lemma. We use this exclusively to prove race-freedom for each pair of adjacent transmissions over some common channel. Consider this pair to be \( S_1 <_{\text{CB}} R_1 \) and \( S_2 <_{\text{CB}} R_2 \). Assume we have already established \( R_1 \triangleleft_{\text{HB}} S_2 \), and that no other transmissions with the same channel are under consideration. From this scenario, we are unable to conclude \( R_1 \triangleleft_{\text{HB}} R_2 \) since the following sequence \( R_2 ; S_1 ; R_1 ; S_2 \) is possible which is consistent \( R_1 \triangleleft_{\text{HB}} S_2 \) but yet violates \( R_1 \triangleleft_{\text{HB}} R_2 \). However, we can conclude \( S_1 \triangleleft_{\text{HB}} S_2 \) since for \( R_1 \triangleleft_{\text{HB}} S_2 \) to hold, it must be the case that \( S_1 \triangleleft_{\text{HB}} R_1 \). By transitivity of \( <_{\text{HB}} \), we can therefore conclude \( S_1 \triangleleft_{\text{HB}} S_2 \).

To reason about these orderings, we propose a constraint ordering system whose syntax is depicted in Fig. 3a. The send and receive events are related using either a HB or a HB relation, while the ordering constraints are composed using either \& or \|. The language supports \( \triangleright^* \) (E) to indicate that an (immutable) event has not occurred yet.

The semantics of the race-free assertions and ordering constraints is given in Fig. 3c, where the proof context is a set of orderings, \( \Pi \), with elements from Fig. 3a. The implication, \( \triangleright^* \) is solved by repeatedly applying the propagation rules ([CB-HB], [HB-HB]) from Fig. 3b.

**Race-free protocol.** For brevity of the presentation, the transmitted messages are ignored in the rest of this subsection, that is, a transmission is described as a tuple \( \text{Role} \times \text{Role} \times \text{Nat} \times \text{Chan} \).
Theorem 3.1 (Race-free protocol). A protocol \( G \) is race-free, denoted by \( RF(G) \), when all the linked transmissions are in a HB relation:

\[
\forall i_1, i_2 \in G \cdot (Adj(i_1, i_2) \Rightarrow i_1 \prec_{HB} i_2)
\]

where \( i_1 \prec_{HB} i_2 \) stands for \( send(i_1) <_{HB} send(i_2) \wedge recv(i_1) <_{HB} recv(i_2) \).

Definition 7 (Race-free adjacent transmissions). Given a protocol \( G \), two adjacent transmissions \( i_1 \in G \) and \( i_2 \in G \) are said to be race-free w.r.t. to each other, denoted by \( RF(i_1, i_2) \), only when they are in a HB relation:

\[
\forall i_1, i_2 \in G \cdot (Adj(i_1, i_2) \Rightarrow i_1 \prec_{HB} i_2)
\]

Proof: Using inductive proving on the definition of linked transmissions and the transitivity of HB. See Appendix, ??.

As mentioned before, we are assuming a hybrid communication system, where both message passing and other explicit synchronization mechanisms are used. Therefore, a simple analysis to check whether the protocol is linear is insufficient. Instead of analyzing the protocol for race channels, we designed an algorithm - **Algorithm 1** - which inserts the race-free constraints as explicit proof obligations within the protocol. We rely on the program verification to detect when these constraints are not satisfied.

**Algorithm 1**: Decorates a protocol with ordering assumptions, and race-free guards

```
input  : G - a global multi-party protocol
output: G' - refined G
1. S ← collect(G)
2. G' ← addGuards(G, S°)
3. G' ← addAssumptions(G', S°)
4. return G'
```

The algorithm is pretty straightforward: it first collects the necessary ordering assumptions and race-free guards, and then inserts them within the global protocol using a simple scheme guided by the unique transmission identifier. The methods used for insertion are not described here for limit of space, but also because they do not represent any theoretical or technical interest. As an intuition though, the insertion follows these principles: \( \oplus(P^{(i)}) \) is inserted immediately after transmission \( i \); \( \ominus(S^{(i)} <_{CB} P^{(i)}) \) is inserted immediately after transmission \( i \); \( \ominus(P^{(i)} <_{HB} P^{(i)}_{1}) \) and \( \ominus(P^{(i)} <_{HB} P^{(i)}_{2}) \) are inserted immediately after transmission \( i \). According to **Algorithm 1** the assumptions are added to the protocol after inserting the guards, therefore, in the refined protocol, the assumptions precede the guards relative to the same transmission.

The collect method derives the ordering relations necessary for the refinement of a user-defined protocol into a race-free protocol. In a top-down approach, collect compiles protocol summaries for each compositional sub-protocol, until each sub-protocol is reduced to a single transmission. It then gradually merges each such summary to obtain the global view of all the orderings within the initial protocol.

**Protocol Summary.** The summary of a protocol \( G \) is a tuple \( S=\langle B, F, A^0, A^\oplus \rangle \), where \( B \) is the left boundary, also called a backtier, mapping roles to their first occurrence and channels to their first transmissions within \( G \), \( F \) is the right boundary, also called frontier, mapping roles to their last occurrence and channels to their last transmissions within \( G \), \( A^0 \) is a set of ordering assumptions which reflect the implicit synchronization ordering relations, and \( A^\oplus \) is a set of ordering proof obligations needed to ensure the race-freedom of \( G \). We denote by \( S^B, S^F, S^0 \) and \( S^\oplus \) the elements of the summary. The \( S^0 \) and \( S^\oplus \) sets are the result of fusing together the communication summaries of each compositional protocol.

**Protocol Boundary.** Each boundary, whether left or right, is a pair of maps \( (K, \Gamma) \), where \( K \) relates protocol roles to events and \( \Gamma \) relates channels to transmission. The fact that a map is not defined for a particular input is indicated by \( \bot \). The constituent elements of these maps are formally described in Fig. 4, where \( \beta^F \) is used to populate the events map - \( K \), and the \( \beta^T \) elements populate the transmissions map - \( \Gamma \). We denote by \( F.K \) and \( F.\Gamma \) the elements of a boundary \( F \).

Finally, the decomposition of the protocol into sub-protocols and the building of summaries is recursively defined as below, where the helper functions, \( h_1 \), \( h_2 \) and \( h_3 \), merge the summaries specific to each protocol composition operator:

\[
\text{collect}(G; G_{2}) \overset{\Delta}{=} h_1(\text{collect}(G_{1}), \text{collect}(G_{2}))
\]
\[
\text{collect}(G_{1} \cdot G_{2}) \overset{\Delta}{=} h_2(\text{collect}(G_{1}), \text{collect}(G_{2}))
\]
\[
\text{collect}(G_{1} \lor G_{2}) \overset{\Delta}{=} h_3(\text{collect}(G_{1}), \text{collect}(G_{2}))
\]
\[
\text{collect}(S^{\ominus} : c) \overset{\Delta}{=} \langle (K, \Gamma), (K, \Gamma), \{ \cap (S^{\ominus} : c) \}, \emptyset \rangle
\]
where \( K=(S^{0} : R^{0} : R^{f}) \) and \( E=(c : S^{\ominus} : R) \)

---

**Figure 4. Elements of the boundary summary.**
Automated Modular Verification for Race-Free Channels with Implicit and Explicit Synchronization

which generates a set of assumptions, where the $\lfloor$ operator (over-loaded over boundaries and maps’ elements) represents the fusion specific to each protocol connector:

$$h_1(S_1, S_2) \equiv (S_1^{st} \lfloor \cdot ]; S_2^{st}, S_1^{te} \lfloor \cdot ]; S_2^{te} \cup \Delta_1 \cup \Delta_2 \cup \Lambda_1 \cup \Lambda_2)$$

where $(\Lambda_1, \Lambda_2) = \text{merge\_adjacent}(S_1^{st}, S_2^{st})$

$$h_2(S_1, S_2) \equiv (S_1^{st} \lfloor \cdot ]; S_2^{st}, S_1^{te} \lfloor \cdot ]; S_2^{te} \cup \Delta_1 \cup \Delta_2 \cup \Lambda_1 \cup \Lambda_2)$$

$$h_3(S_1, S_2) \equiv (S_1^{st} \lfloor \cdot ]; S_2^{st}, S_1^{te} \lfloor \cdot ]; S_2^{te} \cup \Delta_1 \cup \Delta_2 \cup \Lambda_1 \cup \Lambda_2)$$

where $\lfloor \cdot \rfloor$ denotes an update to $K$, such that the value corresponding to $P$ is updated to $P'$, even if $K$ was previously not defined on $P$. Similarly, $\Gamma[c \cdot \beta']$ indicates an update to $\Gamma$ such that $c$ is mapped to $\beta'$. In the case of backtier fusion for sequence, $S_1^{st} \lfloor \cdot ]$ enforces that the resulted backtier reflects all the first transmissions for each channel employed by either $G_1$ or $G_2$, whichever first, and all possible first events for each role in the sequence. Dually, $S_2^{te} \lfloor \cdot ]$, ensures all last transmission and events with respect to the considered sequence will be captured by the newly derived summary.

One of the key points of this phase is captured by the merge of adjacent boundaries, $S_1^{st}$ and $S_2^{st}$, respectively, merge which generates a set of assumptions, $\Lambda_1$ and $\Lambda_2$, and a set of proof obligations, $\Delta_1$. The assumptions capture the happens-before relation between adjacent events, namely the last event of $G_1$ and the first event of $G_2$ with respect to a particular role. Similarly, the guards capture the race-free conditions for all the adjacent transmissions sharing a common channel. Formally this merged is defined by the merge\_adjacent function:

$$\text{merge\_adjacent}(K_1, \Gamma_1), (K_2, \Gamma_2) \equiv (K_1 \cup K_2, \Gamma_1 \cap \Gamma_2)$$

and $\text{Map} = \text{dom}(\text{Map}_1) \cup \text{dom}(\text{Map}_2)$.

In the following, the recursive function merge is overloaded such that it can cater to both event merging (first base-case definition) and transaction merging (second base-case definition):

Local protocol

\[
\mathcal{T} := c \cdot v \cdot \Delta \mid c? v \cdot \Delta
\]

Send/Receive/Transmission

\[
| v \mid \Gamma
\]

Concurrency

\[
| \Gamma \cdot \gamma
\]

Choice

\[
| \gamma \cdot \Gamma
\]

Guard/Assumption

\[
| \oplus(\psi) \mid \oplus(\psi)
\]

(b) Per endpoint

(c) Per channel

4 Local Projection

Based on the communication interface, but also on the verifier’s requirements, the projection of the global protocol to local specifications could go through a couple of automatic projection phases before being used by the verification process. This way, the projection could describe how each party is contributing to the communication, or it could be more granular describing how each communication instrument is used with respect to a communicating party.

**Projection Language.** Fig. 5 describes the two kinds of specification mentioned above. The per party specification language is depicted in Fig. 5a. Here, each send and receive specification name the communication instrument $c$ along with a message $v$ described by a formula $\Delta$. In the per channel specifications, Fig. 5c, the communication instrument is implicit. The congruence of all the compound terms described in Sec. 3 holds for the projected languages as well, with the exception of sequential commutativity since the disjointness conditions for the latter do not hold (eg. either the peer or the channel are implicitly the same for the entire projected specification).

**Automatic projection.** Using different projection granularities should not permit event re-orderings (modulo * composed events).

**Proposition 1 (Projection Fidelity).** The projection to a decomposed specification, such as global protocol to per party,
or per party to per channel, does not alter the communication pattern specified before projection.

To support the above proposition, we have designed a set of structural projection rules, described in Fig. 6. Fig. 6a, describing per party projection rules, is quite self-explanatory, with the exception of the guard projection rule. The latter distinguishes between the projection on the party which needs to prove the guarded happens-before ordering before assuming it, from the party which can soundly assume the ordering without prior proof.

As expected, the per channel projection rules, Fig. 6b, strips the channel information from the per party specifications, since it will be implicitly available. Furthermore, inserting a guard \( \Theta(P^{(i)}) \) between adjacent transmissions on different channels with a common sender ensures that the order of events at the sender’s site is accurately inherited from the corresponding per party specification across different channels. To emphasize this behavior we consider the following sequence of receiving events captured by a per-party specification, \( G \) on:

\[
\begin{align*}
(G)_P: & \equiv c_1 \rhd \cdot A_1; \Theta(P^{(1)}); c_2 \rhd \cdot A_2; \Theta(P^{(2)}); c_2 \rhd \cdot A_3; \Theta(P^{(3)}); c_1 \rhd \cdot A_4 \\
(G)_Pc_1: & \equiv c_1 \rhd \cdot A_1; \Theta(P^{(1)}); \Theta(P^{(2)}); c_2 \rhd \cdot A_3; \Theta(P^{(3)}); c_1 \rhd \cdot A_4 \\
(G)_Pc_2: & \equiv \Theta(P^{(1)}); c_2 \rhd \cdot A_2; \Theta(P^{(2)}); c_2 \rhd \cdot A_3; \Theta(P^{(3)}); c_1 \rhd \cdot A_4
\end{align*}
\]

The above local specification snapshot highlights how local fidelity is secured: the events marked with red boxes are guarded by their immediately preceding events, since they are handled by different channels. A subsequent refinement removes redundant guards, grayed in the example above, since adjacent same channel events need to guard only the last event on the considered channel.

Given the congruence of global protocols and local specifications, the projection is an isomorphism courtesy to the unique labelling and ordering relations carefully inserted after each transmission. Given two protocols \( G_1 \) and \( G_2 \), with \( P_1, P_n \in G_1 \) and \( \neg (\exists P \in G_1: P \notin \{P_1, P_n\}) \), and \( c_1, c_n \in G_2 \) and \( \neg (\exists c \in G_2: c \notin \{c_1, .., c_n\}) \): \( G_1 \equiv G_2 \equiv \{(G_1)_P; j=1..n\} \equiv \{(G_1)_P; j=1..n\} \). This condition holds even for the more granular specifications: \( G_1 \equiv G_2 \equiv \{(G_1)_P_{j,c}; k=1..n\} \equiv \{(G_1)_P_{j,c}; k=1..n\} \). Detailed proofs of the isomorphism under all operators to be provided in the detailed technical report.

And lastly, as discussed in Sec. 2, the communicated-before and happened-before assumptions are projected into a shared store, so that each party can benefit from them (Fig. 6d).

5 Verification of C-like Programs

The user provides the global protocol which is then automatically refined according to the methodology described in Sec. 3. The refined protocol is automatically projected onto a per party specification, followed by a per channel endpoint basis as described in Sec. 4. Using such a modular approach where we provide a specification for each channel endpoint adds natural support for delegation, where a channel (as well as its specification) could be delegated to a third party. These communication specifications are made available in the program abstract state using a combinations of ghost assertions and release lemmas (detailed in the subsequent). The verification could then automatically check whether a certain implementation follows the global protocol, after it had first bound the program elements (threads and channel endpoints) to the logical ones (peers and channels).

Language. Fig. 7 gives the syntax of a core language with support for communication primitives. We omit the details of the Boolean and arithmetic expression and focus on the language support for asynchronous message passing via channels.

Figure 6. Projection rules
Concurrent Separation Logics. Due to its expressive power and elegant proofs, we choose to integrate our session logic on top of concurrent separation logic. Separation logic is an attractive extension of Hoare logic in which assertions are interpreted w.r.t. to some relevant portion of the heap. Spatial conjunction, the core operator of separation logic, $P \land Q$ divides the heap between two disjoint heaps described by assertions $P$ and $Q$, respectively. The main benefit of this approach is the local reasoning: the specifications of a program code need only mention the portion of the resources which it uses, the rest are assumed unchanged.

Fig. 8 defines the state model and the semantics of state assertions. Our approach of layered abstractions permits us to build on top of the traditional storage model for heap manipulating programs with a minimal, yet important extension to account for the race-free assertions. Therefore, we define the program state as the triple comprising a stack $\text{stack}$ and the symbolic heap fragment of separation logic is

$\{\Phi_1, \Phi_2\}$

$\{\Phi_1 \oplus \Phi\}$

$e^*\{\Phi\}$

$e | (c, \{P_1, ..., P_n\}) \mid \text{close}(\hat{c})$

$| \text{send}(\hat{c}, v) | \text{recv}(\hat{c}) | \text{notify}(\text{AI}(E)) | \text{wait}(E)$

$| f(e') | e | e | e | e | \text{if}(b) | e | e \mid e | \text{join}(i_1, i_2)$

**Figure 7. A Core Language**

that the primitives respect their abstract specification, developers could then choose alternative communication libraries, without the need to re-construct the correctness proof of their underlying program.

The verification process follows the traditional forward verification rules, where the pre-conditions are checked for each method call, and if the check succeeds it adds their corresponding postcondition to the poststate. The verification of the method definition starts by assuming its pre-condition as the initial abstract state, and then inspects whether the postcondition holds after progressively checking each of the method’s body instructions.

**Abstract Specification.** We define a set of abstract predicates to support session specification of different granularity. Some of these predicates have been progressively introduced across the paper, but for brevity we have omitted certain details. We resume their presentation here with more details: $\text{party}(P, c', \gamma)$ - associates a local protocol projection $\gamma$ to its corresponding party $P$ and the set of channels $c'$ used by $P$; $\text{gC}(c, P, L)$ - associates an endpoint specification $L$ to its corresponding party $P$ and channel $c$; $\text{Common}(G\#A11)$ - comprises the ordering assumptions shared among all the parties; $\text{Peer}(P)$ - flow-sensitively tracks the executing party, since the execution of parties can either be in parallel or sequentialized; $\text{bind}(P, c')$ - binds a party $P$ to all the channels $c'$ it uses. To cater for each verification phase, the session specifications with the required granularity are made available in the program’s abstract state via the lemmas in Fig. 9.

Channel endpoint creation and closing described by the $[\text{open}]$ and $[\text{close}]$ triples in Fig. 9, have mirrored specification: open associates the specification of a channel $c$ to its corresponding program endpoint $\hat{c}$. close regards the closing of a channel endpoint as safe only when all the parties have finished their communication w.r.t. the closing endpoint.

To support send and receive operations, we decorate the corresponding methods with dual generic specifications. The precondition of $[\text{send}]$ ensures that indeed a send operation is expected. $v \cdot v(v)$, where the message $v$ to be transmitted is described by a higher-order relation over $v$. Should this be confirmed, to ensure memory safety, the verifier also checks whether the program state indeed owns the message to be transmitted and that it adheres to the properties described by the freshly discovered relation, $V(x)$. Dually, $[\text{recv}]$ ensures that the receiving state gains the ownership of the transmitted message. Both specifications guarantee that the transmission is consumed by the expected party, $\text{Peer}(P)$.

The proof obligations generated by this verifier are discharged to a Separation Logic solver in the form of entailment checks, detailed in the subsequent section.

**Entailment.**

Traditionally, the logical entailment between formulas written in the symbolic heap fragment of separation logic is
expressed as follows: $\Delta_\pi \vdash \Delta_c \equiv \Delta_r$, where $\Delta_r$ comprises those residual resources described by $\Delta_\pi$, but not by $\Delta_c$. Intuitively, a valid entailment suggests that the resource models described by $\Delta_\pi$ are sufficient to conclude the availability of those described by $\Delta_c$.

Since the proposed logic is tailored to support reasoning about communication primitives with generic protocol specifications, the entailment should also be able to interpret and instantiate such generic specifications. Therefore we equip the entailment checker to reason about formulae which contain second-order variables. Consequently, the proposed entailment is designed to support the instantiation of such

\[
\begin{aligned}
(s, h, o) \vdash \pi & \iff \\pi = \text{true} \\
(s, h, o) \vdash \forall v (v_1, \ldots, v_n) & \iff \text{struct } d \{ f_1 : \ldots : f_n : f \} \in P \text{ and } h = s(v) \Rightarrow d(f_1 \mapsto v_1, \ldots, f_n \mapsto v_n) \\
(s, h, o) \vdash \Delta_1 \times \Delta_2 & \iff \exists h_1, h_2 : (h_1 \perp h_2 \text{ and } h_1 \not\perp h_2 = h) \text{ and } (s, h_1, o) \vdash \Delta_1 \text{ and } (s, h_2, o) \vdash \Delta_2 \\
(s, h, o) \vdash \Phi_1 \lor \Phi_2 & \iff (s, h, o) \vdash \Phi_1 \text{ or } (s, h, o) \vdash \Phi_2 \\
\text{State } \vdash \text{stack } \times \text{heap } \times \Pi & \iff h_1 \perp h_2 \Rightarrow \text{dom}(h_1) \cap \text{dom}(h_2) = \emptyset \\
\text{Stack } \vdash \text{var } \rightarrow \text{val } \cup \text{loc } \text{heap } \vdash \text{loc } \rightarrow \text{fval } \text{dval} & \iff h = h_1 \not\perp h_2 \Rightarrow \text{dom}(h_1) \cup \text{dom}(h_2) = \text{dom}(h)
\end{aligned}
\]
Automated Modular Verification for Race-Free Channels with Implicit and Explicit Synchronization

Figure 11. Synchronization Primitives for wait-notifyAll variables. However, the instantiation might not be unique, so we collect the candidate instantiations in a set of residual states. The entailment has thus the following form: \( \Delta \rightarrow \{ \text{NOTIFY(E, \( \varnothing \) )} \} \) so we collect the candidate instantiations in a set of residual states. Note that \( S \) is the set of possible residual states. Note that \( S \) is derived and its size should be of at least 1 in order to consider the entailment as valid.

The entailment rules needed to accommodate session reasoning are given in Fig. 10. Other rules used for the manipulation of general resource predicates are adapted from Separation Logic [61].

To note also how \([\text{ENT-RECV}]\) and \([\text{ENT-SEND}]\) are soundly designed to be the dual of each other: while the former checks for covariant subsumption of the communication models, the latter enforces contravariant subsumption since the information should only flow from a stronger constraint towards a weaker one.

Considering the example below, a context expecting to read an integer greater than or equal to 1 could engage a channel designed with a more relaxed specification \((a)\). However, a context expecting to transmit an integer greater than or equal to 1 should only be allowed to engage a more specialized channel, such as one which can solely transmit the exact number 1 \((b)\).

\[
\begin{align*}
(a) & \quad \forall v_1 \geq 1 + \lfloor v_1/v_2 \rfloor v_2 \geq 0 \quad \text{ENT-RECV} \\
& \quad C(c, P, \forall v_1 : v_1 \geq 1 + \forall v_2 : v_2 \geq 0) \quad \text{ENT-CHAN} \\
(b) & \quad \lfloor v_1/v_2 \rfloor v_2 = 1 + v_1 \geq 1 \quad \text{ENT-SEND} \\
& \quad C(c, P, \forall v_1 : v_1 \geq 1 + \forall v_2 : v_2 = 1) \quad \text{ENT-CHAN}
\end{align*}
\]

Soundness. The soundness of our verification rules is defined with respect to the operational semantics by proving progress and preservation. For lack of space we defer the corresponding theorems and proofs to the appendix.

6 Explicit Synchronization

Depending on the communication context, and on the instrument used for communication we could opt amongst a few explicit synchronization mechanisms, such as CountDownLatch, wait-notifyAll, barriers, etc. The choice of these mechanisms are orthogonal to our approach. We have experimented some examples with both CountDownLatch and wait-notifyAll and faced similar results. We chose to formalize the latter since its simplicity suffices for our running example.

The creation of a conditional-variable for wait, \([\text{CREATE}]\) in Fig. 11, releases the specification to verify the calls of notifyAll and wait, respectively, with respect to the session logic orderings.

A call to notify is safe only if the triggering event has occurred already, \([\text{NOTIFY-ALL}]\). In other words, the caller’s state should contain the triggering event information. \([\text{WAIT}]\) on the other hand, releases an ordering assumption conditioned by the send/receive event which is protected by the current wait. The condition has a double meaning in this context: (i) the protected event should have not occurred before a call to wait, and (ii) the ordering is released to the state only after proving that the event indeed occurred, facilitated by the \(\text{Wait lemma}\) in Fig. 11.

The ease of detecting a wait-notifyAll deadlock (within a single synchronization object) is a bonus offered by our logic since it is simply reduced to checking whether there is any context in which a call to \text{wait} terminated without a corresponding \text{notify} call. Formally, this is captured by:

\[
\text{NOTIFY(E, \varnothing) \rightarrow \text{WAIT}(E, \text{emp})} \Rightarrow \perp \text{deadlock}
\]

For more general deadlocks across multiple synchronization objects, we will need to build \(\text{wait-for}\) graphs amongst these objects and detect cycles, where possible, using lemmas similar to the above. For simplicity, these issues are ignored in the current presentation.

7 Modular Protocols

Modularity is essential in designing and implementing new software since it often involves reusing and composing existing components. Software components are modular if their composition in larger software preserves the overall expected computation and, equally important, the safety properties. Checking software compatibility is a known hard problem, even more so when asynchronous communication across components is involved, a heavily used mechanism in building distributed system. In the subsequent subsections, we list a series of extensions to the current logic which support the design of modular protocols. Modular protocols are designed and refined once, and then re-used multiple times to support the design of more complex protocols.

7.1 Labelling.

Thus far our reasoning was based on the fact that each transmission label is unique. To ensure that there is no label clash across all transmission, even in the presence of composed protocols and multiple instantiations of the same protocol, a label is defined as a composition between a parameterized
label root and a unique local id: $i^{#id}$, where such a composed label is a potential label root for nested protocol instantiations.

**Example 1.** Let $H$ be a global protocol which is assembled using protocol $H_0$:

$$H_0(A, B, c, i) \triangleq A^{#1} \rightarrow B : c.$$  

Unrolling $H_0$ within the definition of $H$ results in the following protocol with unique labels:

$$H(A, B, C, c, i) \triangleq A^{#1} \rightarrow B : c; B^{#2} \rightarrow C : c.$$  

### 7.2 Parameterized Frontier for Previous State.

As highlighted in Sec. 3.2 the ordering assumptions and race-free proof obligations are collected via a manipulation of protocol summaries, where a summary consists of a back tier, a frontier and accumulated ordering constraints. The orderings are then added to the global protocol as part of the refinement phase. To extend this approach to composable protocols we equip each protocol with an extra parameter meant to link the current protocol with a generic previous state ensuring that wherever plugged-in, the considered protocol does not create a communication race. Generally speaking, a protocol is described by a predicate in our session logic whose parameters represent the communicating peers, the logical channels used for communication, a root label and the previous state frontier: $H(P', c^*, i, F)$. The same $F$ variable is also used later in Sec 7.4 to denote the frontier for the next stage, via $F'$.

**Example 1 - revisited.** Using the same simple example as in the previous paragraph we highlight the key ideas of parameterized frontier, $F$, as below:

$$H_0(A, B, c, i, F) \triangleq A^{#1} \rightarrow B : c; H_0(B, C, c, #2, F_0).$$

where $F_0 = F[; ]S_1^t$ and $S_1 = \text{CC}(A^{#1} : c)$, therefore $F_0 = F[; ]\{A : (A^{#1}, B, B^{#1})], [c : A^{#1} = B]\}.$

After the refinement phase and the proper instantiation of $H_0$ in the body of $H$, we get:

$$H(A, B, C, i, F) \triangleq A^{#1} \rightarrow B : c; (\oplus(A^{#1} = B); \oplus(K.F(A) <_{\tilde{id}} A^{#1}));$$

$$\oplus(K.F(B) <_{\tilde{id}} B^{#1}); \oplus(K.F(C) <_{\tilde{id}} #1).$$

$$H(A, B, C, i, F) \triangleq A^{#1} \rightarrow B : c; (\oplus(A^{#1} = B); \oplus(K.F(A) <_{\tilde{id}} A^{#1}));$$

$$\oplus(K.F(B) <_{\tilde{id}} B^{#1}); \oplus(K.F(C) <_{\tilde{id}} #1);$$

$$\rho(A^{#1} = B); (\oplus(K.F(A) <_{\tilde{id}} A^{#1}));$$

$$\oplus(K.F(B) <_{\tilde{id}} B^{#1}); \oplus(K.F(C) <_{\tilde{id}} #1));$$

where $\rho = [B/A, C/B, #2/i, F_0/F]$. Knowing $F_0$ and applying the substitution $\rho$, $H$ is then normalized to:

$$H(A, B, C, i, F) \triangleq A^{#1} \rightarrow B : c; (\oplus(A^{#1} = B); (\oplus(F.K(A) <_{\tilde{id}} A^{#1}));$$

$$\oplus(F.K(B) <_{\tilde{id}} B^{#1}); \oplus(F.K(C) <_{\tilde{id}} #1);$$

$$B^{#2\#1} \rightarrow C : c; (\oplus(B^{#2\#1} = C); (\oplus(B^{#1} <_{\tilde{id}} B^{#2\#1}));$$

$$\oplus(F.K(C) <_{\tilde{id}} C^{#1}); \oplus(\#1 <_{\tilde{id}} \#2 \#1).$$

Note that, for the clarity of this example (and subsequent ones), we assume that $F.K(c)$ returns exactly one transmission and $F.K(P)$ returns exactly one event, which is actually the case for this example. However, according to the definition of a map element, Fig. 4, the maps are populated with a composition of transmissions or events. To handle element maps, we should simply form a $<_{\tilde{id}}$ relation for each event transmission returned by the map and add an assumption or a guard for each such created relation.

### 7.3 Sufficient condition for implicit synchronization with the pre-context.

Adding an instantiable frontier to link the current protocol with its usage context enables the refinement process to add the necessary ordering assumptions and guards within the current protocol. As explained in previous sections, these guards hold either when sufficient implicit synchronization is provided, or when explicit synchronization mechanisms are used. To give system designer the option to compose only protocols that are implicitly synchronized with the environment where they are plugged in, we could derive a guard which enforces the implicit synchronization between the protocol and its pre-usage context. To this purpose, we consider a diagramatic view of the ordering relations where each edge is either an HB or CB relation, and aim to find those missing edges which would make the race-free guards (involving the pre-state) hold.

**Definition 8** (Diagrammatic Ordering Relations). A diagramatic view for a set $A^0$ of event orderings is a directed acyclic graph $G(A^0) = (V, Edg)$, where the set $V$ of vertices contains all the events in $A^0$, $V = \bigcup \{V_\Psi | \Psi \in A^0\}$, and the set of edges of $G$ represents all the HB and CB ordering relations in $A^0$, $Edg = \{E_1, E_2 \mid E_1 <_{\tilde{id}} E_2 \in A^0 \lor E_1 <_{\tilde{id}} E_2 \in A^0\}.$

The derivation of the precondition for the implicit synchronization is depicted in **Algorithm 2**: the generic frontier is merged with the backtier of the protocol’s body (line 2) to generate the RF guards which ensure safe composition with the environment. For each such guard $E_1 <_{\tilde{id}} E_2 \in A^0$ to be satisfied, the algorithm searches backwards, starting from $E_2$, a way to connect it to $E_1$ using only ancestor nodes associated to the generic frontier. The ancestors $E(G)$ method returns all the ancestors nodes of node $E$ in graph $G$. The Cartesian product $\times_{\tilde{id}}$ is used to create weak HB relations between $E_1$ and the ancestors of $E_2$. To omit redundancies, the algorithm only considers $E_2$’s earliest ancestors. The Cartesian product between sets of events is defined as follows:

$$S_1 \times_{\tilde{id}} S_2 \triangleq \{E_1 \times_{\tilde{id}} E_2 \mid E_1 \in S_1 \land E_2 \in S_2\}$$

where $E_1 \times_{\tilde{id}} E_2 \triangleq E_1 <_{\tilde{id}} E_2 \lor E_1 = E_2$

The weaker relation $\leq_{\tilde{id}}$ is needed to express that two events might be derived as identical. It is easy to notice that the $\{\text{HB-HB}\}$ propagation rule defined in Fig. 3b can soundly be extended to account for the $\leq_{\tilde{id}}$ relations as well:
Automated Modular Verification for Race-Free Channels with Implicit and Explicit Synchronization

However, the [CB-HB] cannot be fired in the presence of \( \leq_{HB} \) since that would involve changing a \( <_{CB} \) relation into \( <_{HB} \) which would lead to unsoundness. For brevity, we use the shorthand: \( EV(F.K) \) to denote \( \bigcup_{c \in \text{dom}(F.K)} EV(F.K(P)) \) and \( EV(F.I) \) to denote \( \bigcup_{c \in \text{dom}(F.K)} EV(F.I(c)) \).

**Algorithm 2:** Derives the necessary conditions for this protocol to be implicitly synchronized with the pre-context

```
input : H(P', c', i, F) \( \equiv G \) - a global multi-party protocol
output: \( A^\Theta_{pre} \) - a set of orderings guards
1. \( S \leftarrow CC(G) \); \( A^\Theta_{pre} \leftarrow {} \); /* relations between environment and \( H \) */
2. \( (A^\Theta, A^{\Theta'}) \leftarrow F[S]; S^\Theta \); /* relations between environment and \( H \) */
3. foreach \( E_i \in \text{HB} E_2 \in A^\Theta \) do
4. \( A^\Theta' \leftarrow \text{ancestors}(E_2, G(A^\Theta)) \)
   \( \cap \) \( EV(F.K) \cup EV(F.I) \); /* only frontier ancestor */
5. \( A^\Theta' \leftarrow \{ E \mid \text{ancestors}(E, G(A^\Theta)) \cap A^\Theta' \neq {} \} \); /* only earliest ancestors */
6. \( A^{\Theta_{hb}} \leftarrow \{ E \}_X^\Theta \); /* candidate HB relations */
7. \( A^{\Theta_{hb}} \leftarrow \{ E \}_X^\Theta \); /* candidate HB relations */
8. \( A^{\Theta_{pre}} \leftarrow A_{pre}^\Theta \leftarrow \bigcup \{ \forall \delta^{\Theta_{hb}} \in H \}; /* updates the pre guards */
9. end
10. return \( A_{pre}^\Theta \)
```

**Example 1 - revisited.** Using the same simple example as in the previous paragraph, we derive the conditions for which \( H_8 \) could be implicitly synchronized with the body of \( H: \)

\( (A^\Theta, A^{\Theta'}) = ((A^{\#1}, B^{\#1}, A^{\#1}, C^{\#1} <_{CB} B^{\#1}), F(K(A) <_{HB} A^{\#1}, F.K(B) <_{HB} B^{\#1})), \{ \text{send}(F.I(c)) <_{HB} A^{\#1}, \text{recv}(F.I(c)) <_{HB} B^{\#1} \}) \).

Building the graph corresponding to \( A^\Theta \), finding the ancestors of \( A^{\#1} \) and \( B^{\#1} \), and building the HB candidates yields \( A^\Theta_{pre} = \{ \text{send}(F.I(c)) <_{HB} F(K.A) \cap \text{recv}(F.I(c)) <_{HB} F.K(B) \} \) in two iterations of the loop. Given this precondition for \( H_8 \) it is easy to observe with the appropriate instantiation that \( H_8 \) cannot be simply plugged into \( H \) without proper implicit synchronization:

\[ \rho = [B/A, C/B, \#2/i, F_0/F] \]

The above involves proving that \( A^{\#1} <_{HB} B^{\#1} \) and \( B^{\#1} <_{HB} F.K(C) \). This precondition cannot be proved since there is no sufficient implicit synchronization.

Using \( H_8 \) in a different context:

\[ H_1(A, B, C, i, F) \equiv A^{\#1} \rightarrow B : c; H_8(A, B, C, \#2, F_0) \]

yields the following proof obligation:

\[ A^{\#1} <_{HB} A^{\#1} \land B^{\#1} <_{HB} B^{\#1} \]

which is trivially true. Therefore, \( H_8 \) can be plugged in the body of \( H_1 \) without any additional synchronization.

### 7.4 Predicate summary for post-context.

The condition for implicit synchronization ensures race-freedom w.r.t. the pre-usage context. To also ensure communication safety w.r.t. its post-usage context the predicate’s frontier is made available to be plugged-in for merging with the usage site backtrick.

**Example 2.** Using a variation of the previous examples, we highlight the role of a predicate’s frontier at its usage context:

\[ H_2(A, B, C, i, F) \equiv A^{\#1} \rightarrow B : c; B^{\#2} \rightarrow C : c. \]

\[ H_3(A, B, C, i, F) \equiv H_2(A, B, C, i, F) \land B^{\#1} \rightarrow C : c. \]

The frontier of \( H_2 \) is derived to be:

\[ \{(A^{\#1}, B^{\#1}, C^{\#1}(\#2)), [c:i\#2] \} \]

Using this information within the body of the body of \( H_3 \) leads to the following refined protocol:

\[ H_3(A, B, C, i, F) \equiv H_2(A, B, C, i, F) \land B^{\#2} \rightarrow C : c; B^{\#2} \rightarrow C : c. \]

Notice that without the frontier of \( H_2 \), its usage in the body of \( H_3 \) would potentially lead to a race on \( c \) since \( H_3 \) would otherwise not have enough information to add the ordering relations and race-free guards within its body (events labelled \#1\#2).

### 7.5 Recursion.

For recursive predicates we only consider those predicates which are self-contained, that is to say explicit synchronization is only supported within the body of the predicate but not across recursive calls. If this condition is not met, the recursion would have to pass synchronized objects across recursion, creating an unnecessary complicated communication model. To support recursion we: (1) derive the sufficient condition for implicit synchronization with the prestate according to the algorithm described in Sec. 7.3; (2) we derive the summary of the protocol up to the recursive invocation of the considered predicate, as described in Sec. 7.4; (3) given the summary in (2) as ordering context, we next check whether the implicit synchronization condition holds for the recursive invocation. If it does, then the protocol is implicitly synchronized and could be safely plugged-in within any environment which satisfies its implicit synchronization condition.
To check whether this protocol is self-contained, the algorithm first follows the steps described in Sec. 7.3 to derive the sufficient condition for implicit synchronization. This implicit synchronization condition is derived to be:

\[ \text{send}(F.T(c)) \leq_{HB} F.K(A) \land \text{recv}(F.T(c)) \leq_{HB} F.K(B) \]
which, when instantiated to the body of \( H \), it boils down to check that:

\[ A^{(#3)} \leq_{HB} A^{(#3)} \land B^{(#3)} \leq_{HB} B^{(#3)} \]
which is trivially true. Despite containing races within its body, \( H \) is actually safely synchronized across the recursive invocations.

8 Related Work

**Behavioral Types approaches to communication protocols.** The behavioral types specify the expected interaction pattern of communicating entities. Most of seminal works develop type systems on the \( \pi \)-calculus [40, 41] for deadlock [42] and livelock [39] detection. Later developments [43] handle deadlock detection with recursion and arbitrary networks of nodes for asynchronous CCS processes. By using a special recursive model deadlock detection reduce to a check of circularity over dependencies. Linearity is also studied in these systems [58], offering a better reasoning about unbounded dependency chains and recursive types. A generic type system to express types and type environments as abstract processes is proposed in [36]. Following ideas of process algebras, and separation logic [61], the work of [6, 8] introduces a behavioral separation for disciplining the interference of higher-order programs in the presence of concurrency and aliasing. The most intensely studied refinement of the linear behavioral type systems are the session types. Initially proposed for systems with interaction between exactly two peers [34], they have been extended for an arbitrary number of participants [35]. A binary session type describes a protocol from the perspective of only one peer, with support for branching and selection and even with the possibility to delegate the communication to a third party [29, 34]. Other extensions of session types include that of adding support: to handle exceptions across participants [29, 34]; for multithreaded functional languages [30, 50, 53, 57]; for web service description languages [11]; for actor-based languages [28]; for operating systems [26]; for MPI [51]; for C-like languages [56]; for OO languages [22, 23, 31]; and for event-driven systems [44]. Logical foundations of the session types have been studied in the context of a unified theory [14] and linear logic [7, 49, 66]. Session types are able to express effects, and the effect system is powerful enough in order to express session types [57]. In a multiparty session types calculus [35], the user provides a global descriptions of the communication and a projection algorithm computes the local view of each communicating party. Even though it offers general communication channels, it complicates the formalization due to possible race conditions. A less general solution but with a clear formalization to avoid race condition is given in [17]. This restricts the number of channels to the number of distinct communicating participants pairs in the system. Another solution to disambiguate the usage of a single channel across multiple participants is to label the communication actions [9]. Asynchronous semantics is defined either in terms of a projection of the global types [35], or in terms of an asynchronous communication automata [21]. An immediate application of multiparty session types in that of designing deadlock-free choreographies [13], where even if syntactically written in a sequence, certain transmissions which are not in a causal dependence can be swapped. An attempt to describe the content of the exchanged message by adding support for assertions to session types is done in [3]. Our paper focuses on race-freedom [1, 38, 65, 67] over common channels using both implicit and explicit synchronizations [16, 48, 52]. We achieve this abstractly with the help of both \( \leq_{HB} \)-ordering and a novel \( \leq_{CB} \)-ordering. Most of the previous works on session types make the assumption that the underlying system utilizes only implicit synchronization using extra channels, where needed.

**Concurrent Logics for message passing and synchronization mechanisms.** The idea of coupling together the model theory of concurrent separation logic with that of Communicating Sequential Processes [32] is studied in [33]. The processes are modeled by using trace semantics, drawing an analogy between channels and heap cells, and distinguishing between separation in space from separation in time. Our proposal shares the same idea of distinguishing between separation in space and separation in time, by using the \( * \) and \( ; \) operators, respectively. However, their model relies on process algebras, while we propose an expressive logic based on separation logic able to also tackle memory management. Moreover, our communication model is more general, yet more precise, by accounting for explicit synchronizations as well. Heap-Hop [64, 68] is a sound proof system for copyless message passing managed by contracts. The system is integrated within a static analyzer which checks whether messages are safely transmitted. Similar with our proposal, this work is also based on separation logic. As opposed to our proposal, its communication model is limited to bidirectional channels, whereas our model is general enough to also capture buffered channels which may be shared among an arbitrary number of participants. Our proposed session logic is more expressive than previous extensions of separation logic used to express session protocols [18]. It provides more details so that implementation can be safely written. Moreover our logic is specially designed for multiparty protocols. Another extension of separation logic [2] for reasoning about multithreaded programs traces transmitted messages as sets of ordered histories. Our approach does not need to track histories of values, since we rely on an event ordering constraint system which ensures
not only the correct transmission order, but also that of explicit synchronization. The papers [47, 48] propose a verification methodology to prevent deadlocks of concurrent programs communicating via asynchronous shared channels, permit copyless message passing and share memory via locks. This work is similar to ours in the sense that they acknowledge the benefits but also the complications of combining different synchronization mechanisms A resource analysis for π-calculus [63] continues the work of [33] by adding support for two kinds of channels, namely public and private and define two denotational models to reason about both communication safety and liveness. Different concurrent logics based on separation logic have been proposed to reason in a modular manner about the synchronization mechanisms [25, 62]. A permission based logic approach to race-free sharing of heap between concurrent threads is described in [4]. Recently, the authors of [57] use monoids to express and invariants to enforce protocols on shared data in a context of a concurrent separation logic. Numerous static analyses for multiparty protocols have been proposed: deadlock detection [55] by synthesis of a global session graph; minimization of the upper bound of the buffers size [19] by generation explicit new messages for synchronization; computation of recovery strategies in case of failure [34] by solving a causal dependency graph; or liveness and safety checking [46]. Comparing with these analyses, our approach generates separation logic assumptions and proof obligations that are very general, and can moreover be proven in a cooperative manner.

9 Discussions and Final Remarks

Implementation. We have implemented a prototype for Merecurius in OCaml and added support for the communication primitives using the verifier in [15]. As the current work requires specification to be provided, we have only been able to experiment with small examples with our initial experiments. Even so, message passing flawlessly intertwines with different explicit synchronization mechanisms, such as wait – notifyAll and CountDownLatch.

Conclusion. We have designed an expressive multi-party session logic that works seamlessly with both implicit and explicit synchronization to ensure correctness of communication-centric programs. Our approach is built upon first-principle and is based on the use of two fundamental ordering constraints, namely "happens-before" \(\prec_{HB}\) and "communicates-before" \(\prec_{CB}\), to ensure channel race-freedom. Our proof system supports modular verification with the help of automatically generated assumptions and proof obligations from each given global protocol. We have also pioneered the concept of cooperative proving amongst a set of concurrent processes. As part of future work, we intend to provide support for the synthesis of explicit synchronization that can help guarantee race-free communications. We also intend to go beyond the current limits of well-formed disjunctions.

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Figure 12. A Semantic Model of the Core Language

### A Semantics

**Semantic Model.** Following the traditional storage model for heap manipulating programs, the program state is defined as the pair:

\[
\text{State} \equiv \text{Stack} \times \text{Heap}
\]

where a stack \( s \in \text{Stack} \) is a total mapping from local and logical variables \( \text{Var} \) to primitive values \( \text{Val} \) or memory locations \( \text{Loc} \); a heap \( h \in \text{Heap} \) is a finite partial mapping from memory locations to data structures stored in the heap, \( \text{DVal} \):

\[
\text{Stack} \equiv \text{Var} \rightarrow \text{Val} \cup \text{Loc} \quad \text{Heap} \equiv \text{Loc} \rightarrow_{\text{fin}} \text{DVal}
\]

Fig. 12 models the machine for the proposed programming language whose configuration comprises a pair of a *program state* and of *channels states*. The set of channels used by the program are described as a map from a program channel identifier to a *FIFO* list of messages. Moreover, a program is a set of threads identified by a unique id and described by their local state. At any point a thread’s execution influences its local state \( \sigma \), the channels’ state \( \text{CH} \), and advances the program \( \text{P} \).

**Operational Semantics.** The operational semantics is given as a set of reduction rules between machine or thread configurations. Each reduction step is indicated by \( \rightsquigarrow \), where \( \rightsquigarrow^* \) denotes the reflexive and transitive closure of \( \rightsquigarrow \). Moreover, \( \rightsquigarrow^* \) between thread configurations in the machine step, indicates that the proposed semantic is not constrained to a specific scheduler, quantifying over all permissible executions.

Similar to the semantics of [Brookes 2007], a thread reduction could either lead to another thread state and interfere in the states of program channels, \( \langle \langle \sigma, e \rangle, \text{CH} \rangle \rightsquigarrow \langle \langle \sigma', e' \rangle, \text{CH}' \rangle \), or it could signal an error \( \langle \langle \sigma, e \rangle, \text{CH} \rangle \rightsquigarrow \text{error} \). However, a program which is proved correct should never reach an error state. A program terminates in an error-free state if the final configuration reaches the skip expression.

For convenience, only the semantic reduction rules which do not fault are listed down in the subsequent. If any of the premises in these rules do not hold, then the considered reduction leads to error. Moreover, any state which is not captured by the given reduction rules is forced to fault.
A thread configuration TC also accounts for the role played by the instrumented thread.

The current program, an instrumented state expected communication. Besides the local thread state and configuration captures both a map from program channels to heap and stack. The machine configuration is a pair of program channels to semantic model of the program state is instrumented to capture Instrumented Semantic Model.

Small-steps Operational Semantics. As depicted in Fig. 15, the semantic model of the program state is instrumented to capture the communication specification as well as the usual heap and stack. The machine configuration is a pair of program state and channel configuration, where a channel configuration captures both a map from program channels to logical channels, as well as a global protocol describing the expected communication. Besides the local thread state and the current program, an instrumented state \( T_1 \in \text{IState} \) also accounts for the role played by the instrumented thread. A thread configuration \( TC \in T\text{Config} \) is thus defined as:

\[
TC \overset{\text{def}}{=} (T_1, CC)
\]

where \( T_1 = (\sigma, P, e) \) and \( CC = (CS, G) \).

For brevity we implicitly assume the existence of a set of function definitions in the program’s environment.

Small-steps Operational Semantics. Since a machine cannot run such an instrumented semantics, once we prove the soundness of the verifier, we show how the instrumented semantics is correlated to the initial semantics.

The small-step operational semantics are defined by the semantic rules in Fig. 16, Fig. 21 and Fig. 22. These semantic rules are defined using the transition relation \( \rightarrow \) between machine configurations \( \langle PS, CC \rangle \rightarrow \langle PS', CC' \rangle \), and between thread configurations \( \langle T_1, CC \rangle \rightarrow \langle T_1', CC' \rangle \), respectively. Similar to Sec. ??, we use \( \rightarrow^* \) to denote the transitive closure of the transition relation \( \rightarrow \).

Fig. 16 describes the machine reduction rules.

We next distinguish between the possible instantiations of error, namely those reduction faults resulted as a consequence of communication errors:

PROT_ERR indicates that the current reduction refers to a transmission which is not expected within the communication protocol.

RACE_ERR is raised when the \( S_{\text{safe}} \) or \( R_{\text{safe}} \) identify a sending or receiving race condition.
Automated Modular Verification for Race-Free Channels with Implicit and Explicit Synchronization

\[
\begin{align*}
\text{OP-OPEN} & : \quad CH' = CH[\text{res}\rightarrow \text{[]} ] \\
\langle \langle \sigma, \text{open()} \text{ with spec}, \text{CH} \rangle \rangle & \leftrightarrow \langle \langle \sigma, \text{skip}, \text{CH}' \rangle \rangle \\
\text{OP-CLOSE} & : \quad \text{CH} = CH'[\text{\textbar}\rightarrow \text{[]} ] \\
\langle \langle \sigma, \text{close(\textbar)}, \text{CH} \rangle \rangle & \leftrightarrow \langle \langle \sigma, \text{skip}, \text{CH}' \rangle \rangle \\
\text{OP-CLOSE-LEAK} & : \quad \text{CH(\textbar)} \neq \text{[]} \\
\langle \langle \sigma, \text{close(\textbar)}, \text{CH} \rangle \rangle & \leftrightarrow \text{error} \\
\text{OP-SEND} & : \quad \text{chan = CH(\textbar)} \quad \text{chan' := ([}\text{[s]}\text{]e :: chan} \quad \text{CH' := CH[\textbar}\rightarrow \text{chan']} \\
\langle \langle \sigma, \text{send(\textbar, e)}, \text{CH} \rangle \rangle & \leftrightarrow \langle \langle \sigma, \text{skip}, \text{CH}' \rangle \rangle \\
\text{OP-RECV} & : \quad \text{chan = CH(\textbar)} \quad \text{chan = chan' :: [val]} \quad \text{CH' := CH[\textbar}\rightarrow \text{chan']} \\
\sigma = \langle \text{s, h} \rangle \quad \sigma' := \langle \text{s}[\text{res}\rightarrow \text{val}], h \rangle \\
\langle \langle \sigma, \text{recv(\textbar)}, \text{CH} \rangle \rangle & \leftrightarrow \langle \langle \sigma, \text{skip}, \text{CH}' \rangle \rangle \\
\text{OP-RECV-BLOCK} & : \quad \text{chan = CH(\textbar)} \quad \text{chan = [\text{}]} \\
\langle \langle \sigma, \text{recv(\textbar)}, \text{CH} \rangle \rangle & \leftrightarrow \langle \langle \sigma, \text{recv(\textbar)}, \text{CH} \rangle \rangle
\end{align*}
\]

**Figure 14.** Semantic Rules: Per-Thread Reduction

### Machine Config.
- MC ::= PState × Config

### Channels Config
- CConfig ::= CStore × G

### Channel Store
- CStore ::= Endpt → Chan × Role

### Program State
- PState ::= idt → IState

### Instr. Thread State
- IState ::= Stack × Role × P

### Local State
- State ::= Stack × Heap × II

### Thread Config.
- TConfig ::= IState × CConfig

**Figure 15.** An Instrumented Semantic Model of the Core Language

RES_ERR is raised when the communicating thread refers to resources it does not owe.

LEAK_ERR indicates a possible data leak towards unintended recipients.

The expressions which manipulate the channels are given a semantic via the rules in Fig. 18.

To test for protocol conformance we define a predicate, FID(c, P, G) ≜ (safe, m, G'), which checks whether the current transmission is correctly captured within the global protocol. This predicate takes as input a logical channel c which corresponds to the channel which performs the current transmission, a logical party P corresponding to the thread which is performing the transmission and the current state of the communication protocol G. The predicate then encapsulates the safety of the transmission, safe ∈ {SAFE, FAIL} as well as the logical description of the transferred message and an updated communication protocol. To capture that a transmission has been consumed, the corresponding transmission is replaced by a HOLE in the updated global protocol. The FID predicate is inductively defined over the structure of protocol G, as depicted in Fig. 19: the first two cases correspond to a send or receive, respectively, while the third case permits a transmission to overtake the head of the protocol should the current transmission be consumed on a different sender/receiver and channel than expected. The cases for choice and parallel communication emphasize the non-deterministic character of this approach where the current transmission can safely be consumed only by strictly one branch.

To force the machine to fault when it encounters a communication race, we define the predicates Ssafe and Rsafe to check for send or receive race, respectively. The predicates expect a logical peer - corresponding to the role played by the current thread - and a channel specification as arguments. The channel specification is obtained by dynamically projecting the global protocol onto a logical channel as per the projection rules described in Fig. 6a. The holes of the global protocol are ignored during the projection. The predicates are then recursively defined on the structure of the channel specification. Since sending is asynchronous, Ssafe also ensures that the program order and the communication protocol agree on the order of transmissions (P#R in the second case). Ssafe is defined in terms of safe ∈ {SAFE, FAIL}, while Rsafe also captures the blocking behavior of the receive, safe ∈ {SAFE, FAIL, BLOCK}.
Any case which is not explicitly captured by FID, $S_{safe}$ and $R_{safe}$ is deemed to fail.

The communication related expressions are given a semantic via the rules in Fig. 21.

Finally, the $[\text{IOP-ASSERT-PEER}]$ in Fig. 22 updates the configuration of a thread to indicate the role it plays.

\textbf{B Soundness}

This section aims to prove the soundness of the proposed multiparty solution with respect to the given small-step operational semantics by proving progress and preservation. But before proving soundness, it is worth discussing the interference and locality issues addressed by the concurrency formalism approaches in general, and by separation logic based approaches in particular.

\textbf{Interference.} We stress on the fact that, for clarity, the current thesis highlights the effects explicit synchronization has strictly over the communication. The effects that explicit synchronization has over the local heap are orthogonal issue tackled by works such as [Brookes 2007; Feng et al. 2007; Reddy and Reynolds 2012]. This explains the semantic choice for the programming model of Fig. 15 which separates the resources owned by a thread from those owned by

\begin{itemize}
  \item EMPTY($P^*$, $\rightarrow):_R$ \quad $\not\equiv \forall P \in P^* \Rightarrow P \notin \{S,R\}$
  \item EMPTY($P^*$, $Z_1; Z_2$) \quad $\not\equiv$ EMPTY($P^*$, $Z_1$)$\land$EMPTY($P^*$, $Z_2$)
  \item EMPTY($P^*$, $Z_1 \lor Z_2$) \quad $\not\equiv$ EMPTY($P^*$, $Z_1$)$\lor$EMPTY($P^*$, $Z_2$)
\end{itemize}

\begin{figure}[h]
\centering
\includegraphics[width=\textwidth]{figure16}
\caption{Instrumented Semantic Rules: Machine Reduction}
\end{figure}

\begin{figure}[h]
\centering
\includegraphics[width=\textwidth]{figure17}
\caption{Safety Check: Leak-free}
\end{figure}
a communication channel, and where the environment interference only affects the state of communication and not the local state.

**Locality.** Assuming that each message is characterized by a precise formula, each time a thread performs a read it acquires the resource ownership of exactly that heap portion needed to satisfy the formula corresponding to the received message. Similarly, on sending a message the thread releases a resource, i.e. it transfers the ownership of exactly that heap portion determined by the message formula. Since the formulae describing the messages are precise, a transmission can only modify the state of the local heap in one way, releasing or acquiring the resource which is being transmitted, or in other words there is only one possible local transmission which is safe, as highlighted in Fig. 21.

Moreover, a parallel computation $e_1 || e_2$ can be decomposed into the local computations $e_1$ and $e_2$ which are interference-free, except for the communication related interactions carefully guided by a global protocol.

**Definition 9** (Compatible instrumented states). Two instrumented states $\vec{\sigma}_1, \vec{\sigma}_2 \in \text{State}$ are compatible, written $\text{comp}(\vec{\sigma}_1, \vec{\sigma}_2)$, if and only if there exists $P \in \text{Role}$, $e_1, e_2 \in P$, and $\sigma_1, \sigma_2 \in \text{State}$.
Instrumented Semantic Rules: Per-Thread Reduction

Figure 21. Instrumented Semantic Rules: Per-Thread Reduction

Instrumented Semantic Rules: Ghost Transition

Figure 22. Instrumented Semantic Rules: Ghost Transition

Definition 10 (Composition of instrumented states). The composition of two instrumented states \( \overline{\sigma}_1 \) and \( \overline{\sigma}_2 \) in IState, where \( \text{comp}(\overline{\sigma}_1, \overline{\sigma}_2) \) is defined as:

TState such that \( \overline{\sigma}_1 = (\sigma_1, \langle e_1 \rangle) \) and \( \overline{\sigma}_2 = (\sigma_2, \langle e_2 \rangle) \) and \( \sigma_1 \perp \sigma_2 \) and either \( e_1 = \text{skip} \) or \( e_2 = \text{skip} \).
Automated Modular Verification for Race-Free Channels with Implicit and Explicit Synchronization

\[ \sigma_1 \uplus \sigma_1' \models (\sigma_1 \uplus \sigma_2, P, e) \]
with \( \sigma_1 \uplus \sigma_1' \in \text{IState} \).

**Lemma B.1.** For any two instrumented states \( \sigma_1, \sigma_1' \in \text{IState} \), where \( \text{comp}(\sigma_1, \sigma_1') \) and \( \sigma_1 \uplus \sigma_1' \in \text{IState} \), there exists \( CC \in \text{CConfig} \) such that:

1. \( (\sigma_1 \uplus \sigma_1', CC) \models^{\ast} \text{error} \Rightarrow (\sigma_1 \uplus \sigma_1', CC') \models \text{error} \)
2. \( (\sigma_1 \uplus \sigma_1', CC) \models (\sigma_1 \uplus \sigma_1', CC', \langle \langle \sigma_1', CC' \rangle \rangle) \Rightarrow \exists \sigma_1 \uplus \sigma_1' \in \text{IState} \cdot \sigma_1 \models (\sigma_1 \uplus \sigma_1', CC) \models \langle \langle \sigma_1, CC' \rangle \rangle \models \ast \langle \langle \sigma_1, CC' \rangle \rangle \).

**Proof:** Proving (1) is straightforward. Proving (2) requires induction on the length of the derivation and a case analysis on \( e \) which is similar to the standard locality principle of separation logic [Brookes 2007], with the extra judgement on \( CC \).

**Definition 11** (Instrumented Satisfaction). An assertion \( A \) is satisfied in an instrumented thread state \( \sigma_1 \models (\sigma, P, \_ \_ ) \) and channel configuration \( CC \) written \( \sigma_1, CC \models A \), if \( A \) is satisfied in the thread’s local state:

\[ \langle \langle \sigma, P, \_ \_ \rangle, CC \rangle \models A \Rightarrow \sigma \models A \].

**Lemma B.2.** For any two instrumented states \( \sigma_1, \sigma_1' \in \text{IState} \), where \( \text{comp}(\sigma_1, \sigma_1') \) and \( \sigma_1 \uplus \sigma_1' \in \text{IState} \), there exists \( CC \in \text{CConfig} \), and formulae \( \Lambda_1, \Lambda_2 \) such that:

\[ (\sigma_1, CC \models \Lambda_1) \land (\sigma_1', CC \models \Lambda_2) \Rightarrow (\sigma_1 \uplus \sigma_1', CC \models \Lambda_1 \land \Lambda_2) \).

**Proof:** This proof is straightforward from the definition of instrumented states composition and the semantic of *, provided that \( \Lambda_1, \Lambda_2 \) are precise.

The validity of a triple is inductively defined with respect to the reduction rules and satisfaction relation as follows:

**Definition 12** (Validity). A triple \( \{ A_1 \} e \{ A_2 \} \) is valid, written \( \models \{ A_1 \} e \{ A_2 \} \), if:

\[ \forall \sigma \in \text{IState}, CC \in \text{CConfig} : (\sigma_1, CC \models \Lambda_1) \land (\sigma_1', CC \models \Lambda_2) \Rightarrow (\sigma_1 \uplus \sigma_1', CC \models \Lambda_1 \land \Lambda_2) \).

**Theorem 1** (Preservation). For expression \( e \) and states \( \Lambda_1 \) and \( \Lambda_2 \), if \( \{ A_1 \} e \{ A_2 \} \), then \( \{ A_1 \} e \{ A_2 \} \).

**Proof:** We first show that each proof rule is sound: if the premises are valid, then the conclusion is valid. It then follows, by structural induction on \( e \), that every provable formula is valid. We only focus on the communication related rules, since the rest are standard. Assume \( (\sigma_1, CC) \) as the initial configuration for each of the case studies below:

**Parallel Decomposition**

Suppose that \( \sigma_1', CC \models \Lambda_1 \land \Lambda_2 \). Since we only need to prove precise formulae, it results that there exists \( \sigma_1, \sigma_2 \in \text{IState} \) such that \( (\sigma_1 \uplus \sigma_2, CC, \langle \langle \sigma_1', CC \rangle \rangle) \models \text{error} \) or \( (\sigma_1, CC) \models \text{error} \), in which case the whole program faults as per \([\text{JOIN-ERR1}]\) and \([\text{JOIN-ERR2}]\), respectively, or there exists \( \sigma_1 \uplus \sigma_1' \in \text{IState} \) such that \( (\sigma_1', CC, \langle \langle \sigma_1', CC \rangle \rangle) \models \ast \langle \langle \sigma_1', CC \rangle \rangle \) and \( (\sigma_1, CC) \models \ast \langle \langle \sigma_1, CC \rangle \rangle \).
We can conclude therefore that only [op-send-Eprot] can be fired in this case, meaning that: 
\( \langle \sigma', P, \_ \rangle, C' = C(c, P, L) \# Peer(P) \# opened(c, P', \_\_\_\_\_\_\_\_) \wedge P \in P^* \), which trivially holds by (1) and (4) since the only updates to the configuration are represented by the thread state \( \sigma' \), reflecting now the fact that the current thread lost the ownership of the transmitted message, and \( C' \) whose global protocol \( G \) is updated with a HOLE to denote the consumed transmission: \( (P, c) \mid_0 G \equiv L \).

**Receive** If \( \forall \sigma \in \Sigma \), \( \exists \pi \in \Pi \) such that \( \langle \sigma, \pi \rangle \in \text{msg} \), \( C(c, P, \vee \cdot V(v); L) \) also hold: \( (C, G) = CC \land (c, P' = CS(\_\_\_\_\_\_\_\_)) \land (P \in P^*) \). If the corresponding send was not fired yet, the thread will stay in the same state, and when it will be eventually fired, (1) is still satisfied. Since \( C(c, P, \vee \cdot V(v); L) \) (2) specifies a race-free communication w.r.t. channel \( c \) and party \( P \), it follows immediately that the \( (G) \mid_c \) expects for the next reception operation over \( c \) is on \( P \), hence \( \text{SAFE} = R_{\text{safe}}(P, (G) \mid_c) \) (3), assuming the state is not blocked. This ensures that the current execution cannot fault with \text{RACE \_ ERR}, as depicted by [op-recv-Eprot].

We next check whether \( \text{FID} \) holds. (3) implies that the next transmission w.r.t. \( c \) and \( P \) is indeed a receive, we need to check that there are no intermediate transmissions on different channels. From Prop. ??, the projection rules Fig. 6b and (1), that is the communication specification \( C(c, P, \vee \cdot V(v); L) \) expects a receive, it results that there are no other prior unconsumed transmissions expected over \( c \), since that would have involved for the specification to be guarded by \( \Theta(P(\_)) \) prior to the received specified by (2). It then follows that \( \exists \Delta \cdot \text{SAFE} \cdot \vee(\_\_\_\_\_\_\_) = \text{FID}(G, c, P) \) (4), hence this thread cannot fault with \text{PROT\_ERR ([op-recv-Eprot])}. From (2) and (4) we can conclude that \( V(v) = \Delta \). Executing the transmission according to [op-recv-Eprot] results in the update of the current thread state such that it reflects the message ownership gain: \( \exists \sigma_{res} \cdot \sigma' := \sigma \cup \sigma_{res} \) where \( \sigma_{res} \equiv \{ res/v \} \Delta \). Since \( V(v) = \Delta \), then the following also holds: \( \sigma_{res} \equiv V(res) \) (5). With (1), (5) and the update of \( G \) to reflect the consumed receive operation, where \( (P, c) \mid_0 G \equiv L \), we can conclude that the updated thread state satisfies the consequence of \( \text{RECV} \): \( \langle \sigma', P, \_ \rangle, C' = C(c, P, L) \# V(res) \# Peer(P) \# opened(c, P', \_\_\_\_\_\_\_\_) \wedge P \in P^* \).

\( \Box \)

**Theorem 2 (Progress).** If \( \{ \Delta_1 \} \in \text{IState} \) and \( \overrightarrow{\sigma_1} \in \text{IConfig} \cdot \sigma_1, CC \equiv \Delta_1 \) then either \( \exists \sigma \in \text{IState} \) or \( \overrightarrow{\sigma_1} \in \text{IConfig} \cdot \sigma_1, CC \equiv \Delta_2 \).

**Proof:** By induction on the length of the execution and by case analysis on the steps taken we could show that if \( \overrightarrow{\sigma_1}, CC \) is a non-final, fault-free configuration, then \( \overrightarrow{\sigma_1}, CC \) doesn’t get stuck. The communication related cases are straightforward assuming the well-formedness of communication protocols. It may appear as if \text{RECV} could cause a process to get stuck, however, if the protocol which describes the communication is well-formed, as per Def. ??, it is guaranteed for a corresponding sender to get fired within a finite number of machine steps. \( \Box \)
Spec

\[ A \xrightarrow{1} C : c(t_1); \]
\[ A \xrightarrow{2} B : c_2(t_2); \]
\[ B \xrightarrow{3} C : c(t_3) \]

(i) Refinement

\[ A \xrightarrow{1} C : c(t_1); \oplus (A^{(1)}); \oplus (C^{(1)}); \oplus (A^{(1)} \prec_{CB} C^{(1)}) ; \]
\[ A \xrightarrow{2} B : c_2(t_2); \oplus (A^{(2)}); \oplus (B^{(2)}); \oplus (A^{(1)} \prec_{HB} A^{(2)}); \oplus (A^{(2)} \prec_{CB} B^{(2)}); \]
\[ B \xrightarrow{3} C : c(t_3); \oplus (B^{(3)}); \oplus (C^{(3)}); \oplus (B^{(2)} \prec_{HB} B^{(3)}); \oplus (C^{(1)} \prec_{HB} C^{(3)}) \oplus (B^{(3)} \prec_{HB} C^{(3)}) \]

(ii) Per-Party Projection

\[ G#A \triangleq \text{c!}t_1; \oplus (A^{(1)}); \oplus (A^{(2)}); \ominus (1 \prec_{HB} 3)_A \]
\[ G#B \triangleq \text{c?}t_2; \oplus (B^{(2)}); \oplus (C^{(3)}); \ominus (1 \prec_{HB} 3)_B \]
\[ G#C \triangleq \text{c?}t_1; \oplus (C^{(1)}); \oplus (C^{(3)}); \ominus (1 \prec_{HB} 3)_C \]

\[ G#A \equiv \text{c!}t_1; \oplus (A^{(1)}); \oplus (A^{(1)} \prec_{HB} B^{(3)}); \oplus (C^{(1)} \prec_{HB} C^{(3)}) \]
\[ G#B \equiv \text{c?}t_2; \oplus (B^{(2)}); \oplus (B^{(2)} \prec_{HB} B^{(3)}); \]
\[ G#C \equiv \text{c?}t_1; \oplus (C^{(1)}); \oplus (C^{(3)}); \ominus (C^{(1)} \prec_{HB} C^{(3)}) \oplus (A^{(1)} \prec_{HB} B^{(3)}) \]

(ii) Per-Channel Projection

\[ G#A \equiv \text{c!}t_1; \oplus (A^{(1)}); \oplus (A^{(1)} \prec_{HB} B^{(3)}); \oplus (C^{(1)} \prec_{HB} C^{(3)}) \]
\[ G#A \equiv \text{c!}t_1; \oplus (A^{(1)}); \oplus (A^{(1)} \prec_{HB} B^{(3)}); \oplus (C^{(1)} \prec_{HB} C^{(3)}) \]
\[ G#B \equiv \text{c!}t_2; \oplus (B^{(2)}); \oplus (B^{(2)} \prec_{HB} B^{(3)}); \]
\[ G#C \equiv \text{c?}t_1; \oplus (C^{(1)}); \oplus (C^{(3)}); \ominus (C^{(1)} \prec_{HB} C^{(3)}) \oplus (A^{(1)} \prec_{HB} B^{(3)}) \]
\( \text{Spec} \)

\[
\begin{align*}
A & \xrightarrow{1} C : c(t_1); \\
A & \xrightarrow{2} B : c_2(t_2); \\
B & \xrightarrow{3} C : c(t_3)
\end{align*}
\]

(i) Refinement

\[
\begin{align*}
A & \xrightarrow{1} C : c(t_1); \oplus (A^{(1)}); \oplus (C^{(1)}); \oplus (A^{(1)} \prec_{\text{CB}} C^{(1)}); \\
A & \xrightarrow{2} B : c_2(t_2); \oplus (A^{(2)}); \oplus (B^{(2)}); \oplus (A^{(1)} \prec_{\text{HB}} A^{(2)}); \oplus (A^{(2)} \prec_{\text{CB}} B^{(2)}); \\
B & \xrightarrow{3} C : c(t_3); \oplus (B^{(3)}); \oplus (C^{(3)}); \oplus (B^{(2)} \prec_{\text{HB}} B^{(3)}); \oplus (C^{(1)} \prec_{\text{HB}} C^{(3)}); \oplus (B^{(3)} \prec_{\text{CB}} C^{(3)}); \\
\end{align*}
\]

(ii) Per-Party Projection

\[
\begin{align*}
G#A & \triangleq c!t_1; \oplus (A^{(1)}); c_2!t_2; \oplus (B^{(2)}); \ominus (1 \prec_{\text{HB}} 3)^A \\
G#B & \triangleq c_2 ? t_2; \oplus (B^{(2)}); c!t_3; \oplus (B^{(3)}); \ominus (1 \prec_{\text{HB}} 3)^B \\
G#C & \triangleq c ? t_1; \oplus (C^{(1)}); c ? t_3; \oplus (C^{(3)}); \ominus (1 \prec_{\text{HB}} 3)^C \\
G#A1 & \triangleq \oplus (A^{(1)} \prec_{\text{CB}} C^{(1)}); \oplus (A^{(1)} \prec_{\text{HB}} A^{(1)}) \\
G#A & \triangleq \oplus (A^{(2)}); \oplus (B^{(2)}); \oplus (B^{(3)}); \oplus (C^{(1)} \prec_{\text{HB}} C^{(3)}); \oplus (B^{(3)} \prec_{\text{CB}} C^{(3)})
\end{align*}
\]

(ii) Per-Channel Projection

\[
\begin{align*}
G#A & \triangleq c!t_1; \oplus (A^{(1)}); \oplus (A^{(1)} \prec_{\text{HB}} B^{(3)}); \ominus (C^{(1)} \prec_{\text{HB}} C^{(3)}) \\
G#B & \triangleq \ominus (B^{(2)})!t_3; \oplus (B^{(3)}); \ominus (A^{(1)} \prec_{\text{HB}} B^{(3)}); \ominus (C^{(1)} \prec_{\text{HB}} C^{(3)}) \\
G#C & \triangleq \ominus t_1; \oplus (C^{(1)}); \ominus t_3; \oplus (C^{(3)}); \ominus (C^{(1)} \prec_{\text{HB}} C^{(3)}); \ominus (A^{(1)} \prec_{\text{HB}} B^{(3)}) \\
G#A & \triangleq \ominus (A^{(1)}); \ominus t_2 \\
G#B & \triangleq \ominus t_2; \oplus (B^{(2)})
\end{align*}
\]