Abstract
Session types provide static guarantees that concurrent programs respect communication protocols. We give a novel account of recursive session types in the context of GV, a small concurrent extension of the linear $\lambda$-calculus. We extend GV with recursive types and catamorphisms, following the initial algebra semantics of recursion, and show that doing so naturally gives rise to recursive session types. We show that this principled approach to recursion resolves long-standing problems in the treatment of duality for recursive session types.

We characterize the expressiveness of GV concurrency by giving a CPS translation to (non-concurrent) $\lambda$-calculus and proving that reduction in GV is simulated by full reduction in $\lambda$-calculus. This shows that GV remains terminating in the presence of positive recursive types, and that such arguments extend to other extensions of GV, such as polymorphism or non-linear types, by appeal to normalization results for sequential $\lambda$-calculi. We also show that GV remains deadlock free and deterministic in the presence of recursive types.

Finally, we extend CP, a session-typed process calculus based on linear logic, with recursive types, and show that doing so preserves the connection between reduction in GV and cut elimination in CP.

Categories and Subject Descriptors D.3.2 [Language Classifications]: Concurrent, distributed, and parallel languages; D.3.3 [Language Constructs and Features]: Recursion

Keywords Session types, recursion

1. Introduction
Concurrency and communication have become central problems in modern software design and engineering, from hand-held applications relying on remote services to provide key functionality, through traditional applications now running on multi-core hardware, to distributed applications running across data centers. Assuring correct behavior for concurrent programs requires reasoning not just about the type of data communicated, but about the order in which communication takes place. For instance, the messages between an SMTP server and client are all strings representing SMTP commands, but a client that sends the recipient’s address before the sender’s address is in violation of the protocol despite having sent well-formed SMTP commands.

Session types, originally proposed by Honda [22], are an approach to statically verifying communicating concurrent programs. A session type specifies the expected communication along a channel. For example, consider a simplification of the client’s view of the SMTP protocol. After being authenticated, a client has the option of sending one or more messages, each consisting of a sender’s address, recipient’s address, and message, in that order. We could express this with the following session type:

\[
\text{Client} \equiv !\text{FromAddress}.!\text{ToAddress}.!\text{Message}.\text{Client} \oplus !\text{Quit}
\]

where type constants FromAddress, ToAddress, Message, and Quit denote the corresponding SMTP commands. This definition makes use of several session type constructors. The type $!T.S$ denotes sending a value of type $T$ before continuing with behavior $S$, $S \oplus S'$ denotes communicating a choice between behaviors $S$ and $S'$, end denotes the end of a session, and finally we make use of recursive definition to specify repetition in the protocol. A key aspect of session typing is duality. The session type of an SMTP server is dual to that of the client:

\[
\text{Server} \equiv ?\text{FromAddress}.?\text{ToAddress}.?\text{Message}.\text{Server} \oplus ?\text{Quit}
\]

This type definition uses dual features to those in the client’s type: $?T.S$ denotes receiving a value of type $T$ before continuing as $S$ and $S \oplus S'$ denotes receiving a choice between $S$ and $S'$.

We present a novel account of recursive and corecursive session types. Following initial algebra semantics, we characterize recursive computation by catamorphisms (folds) rather than an arbitrary fixed-point operator. Similarly, we characterize corecursive computation by anamorphisms (unfolds). This formulation differs from traditional presentations of recursive session types in three ways. First, we identify dual notions of recursion, corresponding to producers and consumers, rather than having a single self-dual notion of recursion in session types. Second, as they are based on well-founded recursive data types, our recursive session types guarantee termination and freedom from deadlock and livelock. Third, following algebraic ideas of recursion and duality leads to a sound syntactic characterization of duality for recursive session types.

Many previous syntactic formulations of session type duality either incorrectly identify non-dual processes as dual [6, 7] or rely on corecursive expansion of session types [8].

We present our formulation of recursive and corecursive session types as an extension, called $\mu$GV [26], to a core concurrent $\lambda$-calculus called GV. We extend GV with (co)recursive types, (co)recursive session types, and (un)folds over both (co)recursive types and (co)recursive session types, and show that these are suf-
icient to write non-trivial programs using (co)recursive session types. Previous work on GV has minimized the core calculus by encoding session-calculus features in terms of functional features and simple input and output primitives. We continue this thread, showing that recursive session types can be encoded in terms of recursive data types. This result simplifies the concurrent semantics of μGV; for example, it allows us to apply previous results on deadlock freedom and determinism of GV to μGV unchanged. By identifying least and greatest fixed points in the resulting calculus, we obtain a system that admits non-terminating communication, but still guarantees deadlock freedom and productivity.

We also seek to characterize the expressiveness of μGV’s concurrency. To do so, we give a CPS translation from GV into linear λ-calculus (without concurrency), and show that full reduction in the latter simulates reduction in the former. This also allows us to extend standard results on termination in sequential λ-calculus to results for termination of extensions of GV. Most immediately, we can conclude that μGV is terminating. The approach applies equally well to other extensions of GV, such as with polymorphism or non-linear types.

Recent work of Caires and Pfenning [13] and Wadler [34] develops a correspondence between reduction in process calculi and cut elimination in linear logic. GV is closely connected to Wadler’s logic-based process calculus CP: in prior work [26] we give translations between GV and CP such that reduction in each simulates reduction in the other. We extend this observation to (co)recursive session types. We define an extension of CP to include (co)recursive types, following Baele’s formulation of fixed points in classical linear logic [3], and then extend the semantics-preserving translation between CP and GV to include recursive types.

Recent work by Toninho et al. [32] explores corecursive session types from a propositions-as-types perspective. Despite having similar aims, our approach differs from theirs in three significant ways. First, we identify parallels between concurrent and sequential abstractions, in this case between recursive and corecursive data types and recursive session types. Toninho et al., in contrast, develop corecursive session types directly from the corresponding proof theory. This simplifies our concurrent semantics, as we do not have to account for recursive communication directly. Second, we identify two forms of recursive session types—corresponding to encodings based on recursive and corecursive data types—and that their composition can provide unbounded computation. Toninho et al. identify one of these forms, but not the other. Third, as illustrated by the equivalence with μCP, our session types are fundamentally classical, while Toninho et al. build on intuitionistic proof theory.

Thus, for example, our results on the duality of recursive session types do not arise from their approach. We see the coincidence of our typing rules with theirs, despite the significant differences in methodology and foundations, as reinforcing the relevance of both lines of inquiry.

The paper proceeds as follows. We begin with our session-typed functional language, μGV. We present μGV’s functional fragment (§2), extending a core linear λ-calculus with both recursive and corecursive types, and discuss the connection between μGV and languages with non-termination. We present the concurrent fragment of the language (§3), including both recursive and corecursive session types, and compare our approach to traditional characterizations of duality for session types. We give encodings for the concurrent features of μGV in terms of simpler constructs and give a concurrent small-step operational semantics for μGV (§4). We characterize the expressivity of GV concurrency by a CPS translation to a non-concurrent λ-calculus (§5). In doing so, we show that μGV is terminating, and thus (in combination with existing work on GV) free from livelock and deadlock. To establish a strong connection between μGV and linear logic, we present an extension of CP, called μCP, which includes least and greatest fixed points and corresponding recursive and corecursive proof terms, and show semantics-preserving translations from μGV to μCP and vice versa (§6). We conclude by discussing related (§7) and future (§8) work.

2. Functional μGV

We now describe Functional μGV, the functional fragment of μGV. We begin with Functional GV (§2.1), a linear λ-calculus without any form of recursion or iteration. We then discuss two extensions of GV. First, we extend Functional GV with recursive types and cut-elimination in linear logic. GV is closely connected to the calculus. The syntax and typing for recursive types are given at the top of Figure 1. Types include linear implication (T → U), binary and nullary multiplicative products (T ⊗ U and ⊤), additive products (T & U and ⊥), and additive sums (T ⊕ U and T). We write M:N for let () = M in N. The syntax of terms includes (fully applied) constants K.M, used to introduce concurrency; we give the constants and their typing rules in §3.

The typing rules for Functional GV are given in the center of Figure 1; most are standard for linear λ-calculus. The variable rule insists on a singleton environment. Rules for the multiplicative combinators (T → U and T ⊗ U) split their hypotheses, while rules for the additive combinators (T & U and T ⊕ U) duplicate them. In line with its interpretation as falsity, there is no introduction rule for ⊥, nor elimination rule for its dual ⊤.

A small-step operational semantics for Functional GV is given at the bottom of Figure 1. To maintain a close connection with cut-elimination, we define term reduction using weak explicit substitutions [25]. In this approach, we capture substitutions at λ-abstractions instead of applying them directly to the body of the abstraction. Our values thus include closures λx.M, which pair a function abstraction λ.M with a captured environment σ. We extend the typing judgment to include closures by

$$
\frac{\Gamma, x : T \vdash M : U \quad \text{dom}(\sigma) = \text{fs}(M) \setminus \{x\}}{\Gamma \vdash \lambda x. M : T \rightarrow U}
$$

where we write fs(M) to denote the free variables of M. The free variables of a closure are the free variables of the range of σ; capture-avoiding substitution Mσ is defined on the free variables of M. We implicitly treat plain abstractions λx.M as closures λx.(λy.Mσ’), where σ maps each free variable x of M to a fresh variable x’ and σ’ is the inverse of σ.

2.2 Recursion

Next, we extend Functional GV with recursive types and their terms. Our treatment of recursive types is based on the initial algebra semantics of recursion [21]. We extend our language with positive type functors F and their least fixed points μF. We introduce two new term forms: the operator in captures that μF is itself the carrier of an F-algebra, while the fold operator (¬F) captures that μF is initial. This presentation of recursive types has a long history in the functional programming community, dating at least from the treatment of lists in Squiggle [9], and generalized by Meijer et al. [29].

The syntax and typing for recursive types are given at the top of Figure 2. We extend the language of types with type variables X and least fixed points μF, and terms as discussed before. We omit
Syntax

Types

\[ T, U ::= T \rightarrow U \mid T \otimes U \mid 1 \mid T \oplus U \mid 0 \]

\[ | T \land U \mid T \]

Terms

\[ L, M, N ::= x \mid K \mid \lambda x.M \mid MN \]

\[ (M, N) \mid \text{let } (x, y) = M \text{ in } N \]

\[ \emptyset \mid \text{let } () = M \text{ in } N \]

\[ \text{inl} M \mid \text{inr} M \]

\[ \text{case } L \{ \text{inl} x \mapsto M; \text{inr} x \mapsto N \} \]

\[ \text{absurd } M \]

\[ \langle M, N \rangle \mid \text{fst } M \mid \text{snd } M \mid () \]

Typing

\[ x : T \vdash x : T \]

\[ K : T \rightarrow U \quad \Gamma \vdash M : T \]

\[ \Gamma, x : T \vdash M \mid U \quad \Gamma \vdash K : M \mid U \]

\[ \Gamma \vdash \lambda x.M : T \rightarrow U \quad \Gamma, \Gamma' \vdash N : T \]

\[ \Gamma \vdash M : T \quad \Gamma' \vdash N : U \quad \Gamma, \Gamma' \vdash (M, N) : T \otimes U \]

\[ \Gamma, \Gamma' \vdash (M, N) : T \otimes U \quad \Gamma, \Gamma' \vdash \text{let } (x, y) = M \text{ in } N : U \]

\[ \Gamma \vdash M : 1 \quad \Gamma, \Gamma' \vdash N : T \]

\[ \Gamma \vdash \text{let } () = M \text{ in } N : T \]

\[ \Gamma \vdash M : T \quad \Gamma \vdash \text{inl } M : T \oplus U \]

\[ \Gamma \vdash M : T \quad \Gamma \vdash \text{inr } M : T \oplus U \]

\[ \Gamma \vdash \text{case } M \{ \text{inl} x \mapsto N; \text{inr} x \mapsto N' \} : U \]

\[ \Gamma \vdash M : 0 \quad \Gamma \vdash M : T \land U \]

\[ \Gamma \vdash \text{absurd } M : T \]

\[ \Gamma \vdash \text{fst } M : T \land U \quad \Gamma \vdash \text{snd } M : U \]

\[ \Gamma \vdash () : T \]

Values and Contexts

\[ V, W ::= x \mid \lambda x.M \mid (V, W) \mid () \mid \text{inl } V \mid \text{inr } V \mid (V, W) \mid \emptyset \]

\[ \sigma ::= \{ V_1 / x_1, \ldots, V_n / x_n \} \]

where the \( x_i \) are pairwise distinct

\[ E ::= [] \mid KE \mid EDM \mid (E, M) \mid \text{let } (x, y) = E \text{ in } M \]

\[ \text{let } () = E \text{ in } M \]

\[ \text{inl } E \mid \text{inr } E \]

\[ \text{case } E \{ \text{inl} x \mapsto N; \text{inr} x \mapsto N' \} \]

\[ \text{absurd } E \]

\[ \langle E, M \rangle \mid \langle V, E \rangle \mid \text{fst } E \mid \text{snd } E \]

Reduction

\[ \langle \sigma x.M \rangle V \rightarrow_{\gamma} M[V/x \upharpoonright \sigma] \]

\[ \text{let } (x, y) = (V, W) \text{ in } M \rightarrow_{\gamma} M[V/x, W/y] \]

\[ \text{let } () = () \text{ in } M \rightarrow_{\gamma} M \]

\[ \text{case } (\text{inl } V) \{ \text{inl} x \mapsto N; \text{inr} y \mapsto N' \} \rightarrow_{\gamma} N[V/x] \]

\[ \text{fst } (V, W) \rightarrow_{\gamma} V \]

\[ \text{snd } (V, W) \rightarrow_{\gamma} W \]

\[ E[M] \rightarrow_{\gamma} E[M'] \quad \text{if } M \rightarrow_{\gamma} M' \]

The (entirely standard) definition of positivity and well-formedness conditions for types and type operators. As Functional GV is a linear calculus, the typing rule for \( \langle M \rangle \) mandates an empty environment: the evaluation of \( \langle M \rangle \) may require arbitrarily many copies of the expression \( M \), and duplicating \( M \) would require duplicating its assumptions. Functional GV contains only linear assumptions; adding the exponential modality would introduce non-linear assumptions, which could be used in the bodies of catamorphisms.

The semantics of folds is given at the bottom of Figure 2. The extension of values and contexts is unsurprising. Each reduction of a fold \( \langle M \rangle \) amounts to one unrolling of the body \( M \), and depends on the action of \( F \) on terms. With the exception of identity, each type constructor gives rise to both covariant \( F \) and contravariant functors \( F^\ast \). We have the expected typings that if \( M : T \rightarrow U \) then \( F(M) : F(T) \rightarrow F(U) \) and \( F^\ast(M) : F(U) \rightarrow F(T) \).

We use point-free notation for building functors over binary type constructors: for binary type constructor \( \times \), we write \( F \times G \) as shorthand for \( X.F(X) \times G(X) \).

Example. We turn to natural numbers as a characteristic example of recursive types. The definition of the type of naturals parallels the standard (intuitionistic) definition:

\[ \mathbb{N}(X) = 1 \oplus X \]

\[ \text{Nat} = \mu X.1 + X \]

and we can give familiar definitions of the constructors:

\[ zero \equiv \text{in} \( () \) \quad \text{suc}c \equiv \lambda n.\text{in} \(n\) \( \text{in} \)(\text{in} \(n\)) \]

Now consider a standard recursive definition of addition:

\[ 0 + y = y \quad (S(x)) + y = S(x + y) \]

We can rewrite this definition as a curried function of one argument

\[ plus \ 0 = \text{id} \quad plus \ (S(x)) = x + y \]

where \( \text{id} \) is the identity function and \( + \) is function composition. Finally, this version can be expressed using a fold:

\[ plus = \{ \lambda x.\text{case } x \{ \text{inl} (y) \mapsto \text{id}; \text{inr } f \mapsto \text{suc}c o f \} \} \]

The body of the fold has type \( \mathbb{N}(\text{Nat}) \rightarrow (\text{Nat} \rightarrow \text{Nat}) \), so \( \text{plus} \) has type \( \text{Nat} \rightarrow \text{Nat} \rightarrow \text{Nat} \). The product of \( x \) and \( y \), unlike their sum, cannot be computed without duplicating (in some way) either \( x \) or \( y \). In an intuitionistic setting, we might accomplish this by capturing \( x \) in the body of the fold, and using it in each iteration. We cannot do the same in the linear setting. Instead, we will begin by demonstrating terms that duplicate and discard naturals. That is, we show that contraction and weakening are derivable for proposition \( \text{Nat} \). This is an instance of a general result, due to Filinski [18], that contraction and weakening are derivable for the positive constructors in intuitionistic linear logic. In our setting this result is extended to include least fixed points \( \mu F \).

\[ dup = \{ \lambda x.\text{case } x \{ \text{inl} (y) \mapsto (\text{zero}, \text{zero}); \text{inr } z \mapsto (\text{suc}c \text{y}, \text{suc}c z) \} \} \]

\[ drop = \{ \lambda x.\text{case } x \{ \text{inl} () \mapsto () \quad \text{inr } () \mapsto () \} \} \]

We have that \( dup : \text{Nat} \rightarrow (\text{Nat} \times \text{Nat}) \), where the output naturals are equal to the input natural, and \( drop \ : \text{Nat} \rightarrow 1 \), where we can then trivially eliminate the unit value. We can now implement multiplication:

\[ body = \lambda x.\lambda y.\text{case } y \{ \text{inl} () \mapsto (x, 0); \text{inr } y \mapsto \text{let } (x, x') = \text{dup } x \text{ in } \}

\[ (x', \text{plus } y); \quad \text{times} = \lambda x.\lambda y.\text{let } (x', y) = \{ \text{body} \ x \} \text{ in } \text{drop } x'; \text{z} \]

The body of the fold has type \( \mathbb{N}(\text{Nat} \times \text{Nat}) \rightarrow (\text{Nat} \times \text{Nat}) \); in the inductive case, the input pair will be \( (x, y - 1) \), and the result is then \( (x, xy) \). We duplicate \( x \) at each step: one copy is added to the

Figure 1: Functional GV Syntax and Semantics.
As a canonical example of corecursive types, we consider streams of naturals.

\[ S(X) = 1 \& (Nat \times X) \quad Stream = \nu S \]
3. Concurrent $\mu GV$

3.1 Communication and Concurrency

We now consider the concurrent fragment of $\mu GV$. The additional syntax of the concurrent fragment is listed at the top of Figure 4. Our primitive session types include input ($?T.S$), output ($!T.S$), and closed channels (end?). Unlike many session type systems, but in keeping with logically-founded approaches, we have dual types for closed channels rather than a single, self-dual type end. (In our prior work [26] we discuss the semantic and logical consequences of providing a self-dual closed channel.) The remaining features of concurrent $\mu GV$ can be encoded in terms of the primitive concurrent features, and the features of Functional $\mu GV$. These include selection ($S \oplus \overline{S}$), branching ($S \otimes \overline{S}$) and recursive session types ($\mu F$, $\mu^o F$). In traditional session typing notation, $\oplus$ is written as $\oplus$, $\otimes$ as $\&$, $\inl$ as select, and $\case$ as offer. To avoid conflicting with the base features of Functional $\mu GV$ and to emphasize the uniformity of our extensions, we adopt notation that only dualizes the body of session types. We contrast this with the duality between the two forms of recursion. We have corecursive session types where $\mu$ unfolds one iteration of a recursive session type, while $\mu^o$ unfolds one iteration of a recursive session type. Thus, for example, the typing of inl $\overline{M}$ is inverted from the typing of inl, eliminating the session type $S \oplus \overline{S}$.

Our treatment of recursive session types is also guided by initial algebra semantics. We introduce session type variables $\mathcal{X}$ and session functors $\mathcal{F}$. The argument and result of a session functor are both session types. We extend the standard notion of positivity to session functors and ordinary functors of session type (Figure 4). Just as we distinguished between consuming ($! \rightarrow$) and producing (in) values of recursive types, we distinguish between consuming and producing recursive communication. Thus, we have two dual constructors for recursive session types, $\mu F$ for consuming recursive communication and $\mu^o F$ for producing it. The terms inhabiting recursive session types are similar to those for consuming data types; inl $M$ unfolds one iteration of a recursive session type, while $\case M$ consumes a recursive session type; the production and consumption roles are indicated by direction of communication. The typing of inl $M$ reflects its role as an eliminator of session types, parallel to the typing of inl and inr.

As for recursive session types, we have corecursive session types $\nu F$ and $\nu^o F$. The typing rule for out $M$ is a direct reflection of the rule for out $M$. As the communication primitives all consume terms of session type, $\case M$ consumes a channel of type $\nu^o F$, and returns the remaining (empty) expectations of the channel. The differences between recursive and corecursive session types are apparent both in the term formation rules and in the direction of communication: $\case M$ consumes recursive communication, while unfolds $\case M$ produce recursive communication. Though their typing rules appear similar, the differences in communication in inl $M$ and out $M$ are opposite.

The notion of duality is central to session types: if the process holding one end of a channel expects to send a value of some type along that channel, the process holding the other end should expect to receive a value of the same type. The dual $S$ of session type $S$ is defined in Figure 4. The dual of recursive ($\mu F$, $\mu^o F$) and corecursive ($\nu F$, $\nu^o F$) session types is defined in terms of the dualized session function $\overline{F}$. In the definition of $\overline{F}$, note that we not only dualize the body of $F$, but also the variable; this accounts for the duality between the two forms of recursion. We contrast this approach with standard approaches to duality for recursive session types in §3.2.

Promises. While aesthetically appealing, our formation of recursive session types seems somewhat awkward to use. The typing of $\{M\}$ requires that $M$ transform $\overline{F}(S)$ (a session) into $S$ (itself a session); that is, it flattens nested sessions into single sessions. In contrast, most uses of recursive session types transform the data values carried by the session into a result value, only incidentally relying on the nesting of sessions. We can relate these views by observing that $\mu GV$ provides a natural notion of promise, and that promises allow us to treat arbitrary types as session types. Promises [28] introduce asynchrony between the computation of a value and its use; a promise of type $T$ denotes a value of type $T$ which may not yet have been computed. A promise of type $T$ arises naturally in our setting as a channel of type $?T.\text{end}$. We write $?T$ to abbreviate $?T.\text{end}$, and define mappings between promises and values:

- $\text{un}^\circ : ?T \rightarrow T$
- $\text{un}^\circ M = \text{let } (c,z) = \text{receive } M \text{ in } \text{wait } c;z$
- $\text{en}^\circ : T \rightarrow ?T$
- $\text{en}^\circ M = \text{fork } (\lambda x. \text{send } (M,x))$

Figure 4: Concurrent $\mu GV$ Terms and Typing

<table>
<thead>
<tr>
<th>Syntax</th>
</tr>
</thead>
<tbody>
<tr>
<td><strong>Session types</strong></td>
</tr>
<tr>
<td>$S ::= \top T.S \mid ?T.S \mid \text{end}_1 \mid \text{end}_2$</td>
</tr>
<tr>
<td>$S::= S \oplus \overline{S} \mid S \otimes \overline{S} \mid 0 \mid X \mid \overline{X} \mid \mu F \mid \nu F \mid \nu^o F \mid \nu^o F$</td>
</tr>
<tr>
<td><strong>Types</strong></td>
</tr>
<tr>
<td>$T, U ::= \cdots \mid S$</td>
</tr>
<tr>
<td><strong>Session functors</strong></td>
</tr>
<tr>
<td>$F ::= \overline{\mathcal{X}}$</td>
</tr>
<tr>
<td><strong>Terms</strong></td>
</tr>
<tr>
<td>$L, M, N ::= \cdots \mid \text{inl} M \mid \text{inr} M$</td>
</tr>
<tr>
<td>$\text{case } L { \text{inl} x \rightarrow M; \text{inr} y \rightarrow N }$</td>
</tr>
<tr>
<td>${M} \mid \text{inl} M \mid \text{out} M \mid \overline{M}$</td>
</tr>
<tr>
<td><strong>Constants</strong></td>
</tr>
<tr>
<td>$K ::= \text{send} \mid \text{receive} \mid \text{fork} \mid \text{wait} \mid \text{link}$</td>
</tr>
<tr>
<td><strong>Duality</strong></td>
</tr>
<tr>
<td>$\overline{\top S} = ?T.S$</td>
</tr>
<tr>
<td>$\overline{?! T.S} = T.S$</td>
</tr>
<tr>
<td>$\overline{S \oplus \overline{S}} = S \otimes \overline{S}$</td>
</tr>
<tr>
<td>$\overline{S \otimes \overline{S}} = S \oplus \overline{S}$</td>
</tr>
<tr>
<td>$\overline{\overline{S}} = S$</td>
</tr>
<tr>
<td>$\overline{\overline{\mathcal{X}}} = \mathcal{X}$</td>
</tr>
<tr>
<td>$\overline{\overline{\mathcal{T}}} = \mathcal{T}$</td>
</tr>
<tr>
<td><strong>Positivity</strong></td>
</tr>
<tr>
<td>$\xi \in { X, \mathcal{X} }$</td>
</tr>
<tr>
<td>$\text{p}(\xi, ?T.S) = \text{p}(\xi, T) \land \text{p}(\xi, S)$, where $p \in { \text{neg}, \text{pos} }$</td>
</tr>
<tr>
<td>$\text{pos}(\xi, !T.S) = \text{neg}(\xi, T) \land \text{neg}(\xi, S)$</td>
</tr>
</tbody>
</table>

| **Typing** |
| $\Gamma \vdash M : S \oplus \overline{S}$, $\text{send} : T \rightarrow !T.S \rightarrow S$ |
| $\Gamma \vdash \text{inl} M : S \rightarrow T.S \rightarrow S$ |
| $\Gamma \vdash \text{fork} : (S \rightarrow \text{end}_1) \rightarrow \overline{S}$ |
| $\Gamma \vdash M : S \otimes \overline{S}$, $\text{wait} : \text{end}_2 \rightarrow 1$ |
| $\Gamma \vdash \text{inr} M : S \rightarrow \text{end}_1$ |
| $\Gamma \vdash L : S \oplus \overline{S}$, $\Gamma, x : S \vdash M : T$, $\Gamma, x : S' \vdash N : T$ |
| $\Gamma \vdash \text{case } L \{ \text{inl} x \rightarrow M; \text{inr} x \rightarrow N \} : T$ |
| $\Gamma \vdash L : 0 \rightarrow ?T$ |
| $\Gamma \vdash M : \mu F \rightarrow M : \overline{F}(S) \rightarrow S$ |
| $\Gamma \vdash \text{case } L \{ \} : ?T$ |
| $\Gamma \vdash \text{inl} M : \overline{F}(\mu F) \rightarrow \{M\} : \mu F \rightarrow S$ |
| $\overline{\mu F} : \nu F \rightarrow \nu F$ |
| $\overline{\nu F} : \nu^o F \rightarrow \nu^o F$ |

(Shaded terms and types, and their typing and duality, can be encoded in terms of the remaining terms and types.)
The operation \( un^\mu M \) retrieves a value from promise \( M \) (blocking until it is available), while \( en^\mu M \) constructs a new promise, already containing the value of \( M \). Note that \( un^\mu \circ en^\nu \) and \( en^\nu \circ un^\mu \) are both observationally equivalent to the identity function (the second is a trivial example of channel forwarding). The dual of the type \(?T \rightarrow !T\) is the type \(!T \rightarrow ?T\), which we will abbreviate \(!T\). We can also define an operation to introduce and eliminate channels of type \(?T\) (that is, to provide results to unfulfilled promises):

\[
un^\mu : (S \rightarrow T) \otimes !T \rightarrow S \\
en^\nu : S \rightarrow !S \\
\]

The appearance of the continuation type \( S \) may be surprising; this is a consequence of the different treatment of the continuation in send and receive. As \( \mu \textit{GV} \) lacks polymorphism, we will write \( un^\mu M \) to denote the substitution of \( M \) into the definition of \( un^\mu \), rather than to denote an application in \( \mu \textit{GV} \), and similarly for the other definitions in this section.

**Example: Recursive Sessions.** We now present several examples of channels of naturals, building on our earlier representation of naturals numbers. We introduce type abbreviations for such channels:

\[
NC(X) = \text{end} \uplus \ ?Nat.X \\
\text{Nats} = (\mu \textit{NC})
\]

We begin with a term that sends two naturals along a given channel:

\[
twoNats : \text{Nats} \rightarrow \text{end}
\]

\[
twoNats = \lambda c. \text{let } c = \text{send (zero, inr (in \ c)) in} \\
\text{let } c = \text{send (succ zero, inr (inl c)) in} \\
\text{inr (inl c)}
\]

Now we present a slightly more interesting example. Given some starting natural \( n \), we send the sequence \( n, n+1, \ldots \) along a channel. We rely on \( \text{Nat} \) itself being defined recursively.

\[
downFrom : \text{Nats} \rightarrow \text{end}
\]

\[
downFrom n = \text{let } (n', k) = (\text{body}) n \text{ in drop } n' \langle k \rangle
\]

\[
\text{body} : N(\text{Nat} \otimes (\text{Nats} \rightarrow \text{end})) \rightarrow (\text{Nat} \otimes (\text{Nats} \rightarrow \text{end}))
\]

\[
\text{body} z = \text{case } z [\text{inl } (\lambda y, k) \rightarrow z] \\
\text{inr (y, k)} \rightarrow \text{sc y k}
\]

\[
z c : \text{Nat} \otimes (\text{Nats} \rightarrow \text{end})
\]

\[
z c = (\text{zero, } \lambda c. \text{let } c = \text{send (zero, inr (in \ c)) in} \\
\text{inr (inl c))}
\]

\[
\text{sc} : \text{Nat} \otimes (\text{Nats} \rightarrow \text{end}) \rightarrow (\text{Nat} \otimes (\text{Nats} \rightarrow \text{end}))
\]

\[
\text{sc y k} = \text{let } (y, y') = \text{dup y in} \\
\text{succ y, } \lambda c. \text{k (send (y', inr (inl c))))}
\]

We have a similar challenge in defining \( \text{body} \) as we did in defining \( \text{times} \): at each step of the recursion, we must both send a value along the channel and produce the same value for the next step. Observe that \( \text{body} \) has type \( N(\text{Nat} \otimes (\text{Nats} \rightarrow \text{end})) \rightarrow (\text{Nat} \otimes (\text{Nats} \rightarrow \text{end})) \), computing both the next natural in the sequence and the function that sends it along a channel.

We can also write functions that consume channels of naturals. For a simple example, we could compute the sum of the naturals received along a channel.

\[
\sum : \text{Nats} \rightarrow \text{Nat}
\]

\[
\sum = \lambda \text{c. case } c [\text{inl } \rightarrow \text{wait } c; \text{en' zero} \\
\text{inr } c \rightarrow \text{let } (x, c) = \text{receive } c \in \\
\text{let } y = \text{un' c in} \\
\text{en' (plus x y))}]
\]

We wrap the running sum in a promise to lift it to session type: the body of the fold has type \( NC(\text{Nat}) \rightarrow \text{Nat} \), so \( \text{sum} \) has type \( \mu \text{NC} \rightarrow ?\text{Nat} \). We could compose this with one of the producers above to compute a value, such as:

\[
\text{un' (sum (fork (downFrom 1)))}
\]

where we write \( ? \) to indicate the representation of natural \( n \). This term will evaluate to \( 10 \).

We can also define channel transformers. For example, we could compute the running total of the stream, inserting the total (to that point) after each element.

\[
\text{running} : \text{Nat} \rightarrow \text{Nat} \rightarrow \text{end}
\]

\[
\text{running } c = \text{un' ((}\lambda c \rightarrow \text{done } c; \text{inr } c \rightarrow \text{more } c))\text{ c 0}
\]

\[
\text{done } c = \text{en' (}\lambda c. \text{let } c = \text{receive } c \in \\
\text{let } k = \text{un' c in} \\
\text{let } (y, y') = \text{dup y in} \\
\text{en' (}\lambda c. \text{let } (w, w') = \text{dup (plus y' c)) in} \\
\text{let } d = \text{send (y, inr (inl d)) in} \\
\text{let } d = \text{send (w, inl (in' d)) in} \\
\text{w k d}
\]

Note that the body of the fold has type \( NC(\text{Nat} \rightarrow \text{Nat} \rightarrow \text{end}) \rightarrow ?(\text{Nat} \rightarrow ?\text{Nat} \rightarrow \text{end}) \); the \( \text{Nat} \) argument stores the running sum, and so is initialized to zero by \( \text{running} \).

**Example: Corecursive Sessions.** The treatment of corecursive session types is similar to that of recursive session types. Nevertheless, we give a short example of their use. The previous example showed a term \( \text{downFrom} \) that sent all the naturals below a starting value along a channel. With corecursive session types, we can give a term that (potentially) sends all the naturals above a starting value along a channel. We begin with the terms; unlike the previous version, it is now the side sending naturals that offers to choose to continue:

\[
NC'(X) = \text{end} \uplus \ ?Nat.X \\
\text{Nat}' = (\nu' \text{NC'})
\]

We can now define a process that sends an increasing stream of naturals:

\[
\text{body} : \lambda n. \lambda c. \text{case } c [\text{inl } \rightarrow \text{send m }; \text{inr } c \rightarrow \text{let } (m, m') = \text{dup m in} \\
\text{let } c = \text{send (m, c) in} \\
\text{send (succ m', c)}
\]

\[
\text{upFrom} = (\\text{body})
\]

As before, we rely on duplication and discard being defined for natural numbers. We have that \( \text{body} : \text{Nat} \rightarrow NC(\text{Nat} \rightarrow \text{end}) \), and so \( \text{upFrom} : \text{Nat} \rightarrow \text{Nat}' \rightarrow \text{end} \). A simple consumer of such a channel might compute the sum of the first two values on the channel:

\[
\text{firstTwo } c = \text{let } (x, c) = \text{receive (inr (out' c)) in} \\
\text{let } (y, c) = \text{receive (inr' (out' c)) in} \\
\text{wait (inl (out' c)); x + y}
\]

### 3.2 Recursion, Duality, and Session Types

We relate our statement of duality to previous accounts of recursive session types. Most existing approaches use equirecursive, self-dual recursive session types, and do not include dualized type variables. Our system lacks self-dual constructs, and so cannot encode self-dual recursive types. However, we can imagine extending our language with a construct \( \mu^\nu F = \mu^\nu F \) such that \( \mu^\nu G = \mu^\nu G \).

Honda et al. [23] originally proposed recursive session types for a system of first order session types in which messages did not
include channel names. Duality was given by μ²X.T = μ²X.T, where X = X. To distinguish this notion from ours, we write naive(T) instead of T. In contrast, our approach gives μ²X.T = μ²X.(T/X). It is not hard to see that logical duality coincides with naive duality for first-order session types. Intuitively, if μ²X.T is first-order, then if we compute μ²X.(T/X) each instance of X in T will first be dualised by T and then again by the substitution (T/X).

Independently, Bono and Padovani [10] and Bernardi and Hennessy [6] observe that naive duality is not enough for higher-order session types, that is, session types with support for delegation. Consider S = μ²X.μ²Y.X.X, the logical dual of S is μ²X.(μ²Y.X.X), whereas the naive dual of S is μ²X.(μ²Y.X.X), which is (equircursively) equivalent to μ²X.(μ²Y.X.X). It is not difficult to show that the logical dual yields the correct behaviour, whereas the naive dual does not. They (each) proposed a new definition of duality for recursive session types, using a selective form of substitution which applies only inside carried types. Later, Bernardi and Hennessy [7] observed that even this approach fails on examples such as μ²X.μ²Y.Y.X, Y. They propose converting each recursive session type into a so-called m-closed recursive session type before applying native duality. A recursive session type μ²X.T is m-closed if X does not occur free inside a carried type in T. It is straightforward to show that every recursive session type is (equircursively) equivalent to an m-closed one, and that, as they they are essentially first-order, naive duality and logical duality coincide on m-closed recursive session types.

Duality for recursive session types clearly needs to be treated carefully. We are encouraged that our definition coincides with the state of the art for equircursive self-dual session types. We believe that this also shows the value of our deconstruction of recursive session types into well-understood primitives: we are guided immediately to a correct, compact, and general definition of duality.

Remark. Dualized session type variables are redundant in equircursive session types, as every session type μ²X.T is equivalent to μ²X.T μ²Y.T[X/T]/X, where Y is a fresh type variable. However, dualized session type variables do yield a cleaner compositional definition of duality.

4. Communication with Concurrency

4.1 Encoding Concurrent Features

Our prior work on GV [26] focuses on keeping the core language as simple as possible. For example, rather than include branching and choice in the concurrent semantics directly, a choice can be encoded as the promise of a (data type) sum. Kobayashi et al [24] and Dardha et al. [17] make similar use of linear promises to relate data types and session types in π-calculi. We extend this view to include recursive session types. There are two challenges in doing so: a) we must encode session functors and their use of session type variables, and b) we must encode their fixed points.

Our translation is given by the homomorphic extension of the rules in Figure 5. We underline those portions of the translation that introduce purely administrative reduction. The session functor F is translated to the ordinary function F, using promises to lift an ordinary type variable to session type. This approach naturally accounts for the use of dualized session type variables; for example, if F(X) = [X,X], then we have that F′(X) = !(X)!X = !(X)!X. Recursive and corecursive session types are interpreted as promises of recursive and corecursive data types, and the interpretation of their terms is directed by the interpretation of their types. The definitions of (M) and [M] may seem surprisingly complicated. In fact, we can present a different form of session-typed catamorphisms and anamorphisms, directly encoded in terms of recursive types:

\[ (M)^S N = (M)(un^S M) \]
\[ [L]^M N = send([\lambda x. fork (\lambda c. L x c)] M, N) \]

with the following typing rule, which restricts the translation to session functors for a direct use of their encoding:

\[ \vdash M : F_a (\neg T) \rightarrow T \]
\[ \vdash \phi : \mu \nu F_a \rightarrow T \]

We can see that the encoding of (\neg \phi)^S is an instance of (\neg \phi)^S, that the encoding of (\neg \phi)^S is an instance of (\neg \phi)^S, and that our examples can be written directly using the alternative forms (removing calls to en as necessary). Nevertheless, we have preferred (\neg \phi)^S and (\neg \phi)^S as they do not rely on details of our encoding and are closer to the algebraic intuition.

4.2 Concurrent Semantics

We give a concurrent semantics of μGV, building on the small-step operational semantics for Functional μGV given in the last section. As recursive session types and their terms can be encoded in terms of the core concurrency features, our semantics is mostly unchanged from that of our prior work [26]; Figures 6 and 7 give the syntax and typing of configurations and configuration contexts; we will write \( \Gamma \vdash C \vdash T \) to denote that there is some \( \phi \) such that \( \Gamma \vdash \phi \vdash C \vdash T \). Figure 8 gives reductions and configuration equivalence. Because of the importance of promises in our interpretation of μGV’s concurrent features, we give special cases of the functor map for the promise functor. These have the same behavior as that given in the general case, but expose potentially administrative reductions sooner. Our treatment of link repair is a defect in our prior account of GV and restores the diamond property for GV’s concurrent semantics.

The concurrent semantics of μGV is defined by a reduction relation on collections of parallel threads, called configurations. The syntax of μGV configurations is given in Figure 6, and includes threads, name restriction and composition of configurations. Be-
Therefore, ignoring all internal reductions would trivially identify all processes that compute the same results. We intend a finer characterisation, in which we ignore only administrative reductions. We have already identified (by underlining) the relevant sources of administrative reductions. We say that a reduction is administrative (→a) if all the reduced subexpressions are underlined. For example, the reduction of $H[send(x, y)] \rightarrow H'[receive(x)]$ to $H[x]$ is administrative, but the reduction of $H[send(M, x)]$ to $H[send(M', x)]$ is not (unless $M$ is itself identified as administrative). We write $\rightarrow^a$ for the reflexive, transitive closure of $\rightarrow_a$, and write $C \leadsto C'$ to denote $C \rightarrow^a C'$ for all $C, C'$. Finally, we can adapt the standard notion of weak bisimulation to our setting.

**Definition 4.** A relation $\mathcal{R}$ on configurations is an administrative weak bisimulation if, for each $C_1 \mathcal{R} C_2$, whenever $C_1 \rightarrow C_3$, then there is a $C_2'$ such that $C_2 \mathcal{R} C_2'$ and $C_1 \mathcal{R} C_3$ and similarly for $C_2$. We define administrative weak bisimilarity $\approx_a$ to be the union of all administrative weak bisimulations.

We can now relate the encoding of recursive session types to their expected semantics:

**Theorem 5.**

1. If $\Gamma \vdash M : \mathcal{F}(S) \rightarrow S$, then
   $$(\mu x)(H[\mu y] x) \parallel (\mathcal{F}[\mu y] x) \rightarrow H'[x]$$

2. If $\Gamma \vdash M : S \rightarrow \mathcal{F}(S)$, then
   $$(\mu x)(H[\mu y] x) \parallel (\mathcal{F}[\mu y] x) \rightarrow H'[x]$$

The key observations to establishing this result are that $un^i$ and $en^i$ introduce only incidental additional concurrency.

**Lemma 6.**

1. If $\Gamma \vdash M : \mathcal{F}(un^i) (\mathcal{F}(\mu y) x)$ then $M \approx_a (\mu x)(H[\mu y] x) \parallel H'[x]$

2. $E[\mathcal{F}(un^i) (\lambda x.e)(M x) N] \approx_a E[\mathcal{F}(M) N]$

3. $E[\mu x.(\lambda y.e)(\mu y)(\lambda x.e)(\mu y)(x)] \approx_a E[\mathcal{F}(M) N]$

The first is entirely straightforward, the second and third can be shown by induction on the structure of $\mathcal{F}$. The theorem follows directly from the lemmas and the definition of reduction.

5. Communication without Concurrency

We now show, via a CPS translation, that reduction in $\mu GV$ can be simulated by reduction in Functional $\mu GV$. We begin with a standard left-to-right call-by-value CPS translation from the core calculus into itself (Figures 9 and 10), where $R$ is a fixed return type. In the rest of this subsection, we extend the CPS translation to session types and show that the CPS translation preserves reduction. As a corollary, we obtain that $\mu GV$ is strongly normalising.

Following Danvy and Nielsen [15], we can mechanically transform the naive CPS translation $\mathcal{N}[-]$ of Figure 10 into a compositional first-order one-pass CPS transformation $\mathcal{K}[-]$. By carefully distinguishing between values and non-values, the one-pass translation ensures that (most) administrative redexes are contracted by the translation itself. Contracting these redexes enables a tight simulation result (Theorem 10). Due to lack of space, we omit the (entirely standard) details of the one-pass variant of the translation.
transformation of configurations, we restrict attention to a canonical
class of configurations. We write \(C_1 \parallel C_2\) to denote a parallel
composition in which channel \(x\) has input session type in \(C_1\) and
the dual output session type in \(C_2\). We say that a configuration \(C\)
is well-oriented (WO(\(C\))) if all of the parallel compositions in \(C\)
are of this form, and in any link configuration \(z = x \leftrightarrow y\) in \(C\),
\(x\) has input session type. Without loss of generality, we need only
consider reduction on well-oriented compositions.

**Lemma 7.** If \(\Gamma \vdash C : T\), then there exists well-oriented \(C' \equiv C\).

The translation of the main thread is the only place the continuation
is actually used. The translation of a child thread supplies the identity
continuation, which is well-typed as child threads always have type end.
Name restrictions themselves are ignored, but names are used in the
translation of well-oriented parallel composition. It is straightforward to verify that the CPS translation preserves typing.

**Theorem 8 (Type soundness).**

1. If \(\Gamma \vdash M : T\), then \(\mathcal{K}[\Gamma] \equiv \mathcal{K}[M] : (\mathcal{K}[T] \rightarrow R) \rightarrow R\).
2. If \(\Gamma \vdash C : T\) and \(C \) is well-oriented, then \(\mathcal{K}[\Gamma] \vdash \mathcal{K}[C] : (\mathcal{K}[T] \rightarrow R) \rightarrow R\).

To reason by induction over the reduction rules, which are
defined in terms of evaluation contexts and configuration contexts,
we extend the CPS translation to contexts (Figure 12). Evaluation
contexts are interpreted as functions. The CPS translation of a
configuration context takes two arguments. The first argument is a
meta-level function, which we instantiate with the CPS translation of
an appropriate configuration. The second is a continuation. The
CPS translation respects decomposition of contexts.

**Lemma 9.**

1. If \(\Gamma \not\vdash \_\) and \(\Gamma \vdash E[T] : T\), then \(\mathcal{K}[\Gamma][E[T]]k = \mathcal{I}[E[T]](\lambda x.E[E]k)\).
2. If \(\Gamma \vdash D[C] \) and \(D[C]\) is well-oriented, then \(\mathcal{K}[\Gamma][D[C]]k = \mathcal{D}[D][\mathcal{K}[C]]k\).

To simulate all reduction paths in \(\mu GV\) by reduction in Functional \(\mu GV\), we must allow reduction under lambda abstractions.

Otherwise, we would be limited to a single schedule in which order
of communication is determined by the outermost input communication.
For an intuition of this schedule, consider the translation of \(C \parallel C'\). Reduction in \(C\) cannot proceed until \(C\) is ready to receive a
result along \(x\), even if \(C'\) could perform some internal communication.
Allowing reduction under lambda abstractions allows all
valid schedules to be simulated. We define \(M \rightsquigarrow N\) by:

\[
\begin{align*}
M & \rightsquigarrow N \\
\lambda x.M & \rightsquigarrow \lambda x.N \\
M \parallel N & \rightsquigarrow \lambda x.N \parallel \lambda x.M
\end{align*}
\]

The following theorem states that the CPS translation simulates
\(\mu GV\) reduction.

**Theorem 10 (Simulation).**

1. If \(\Gamma \vdash M : T\) and \(M \rightsquigarrow N\), then \(\mathcal{K}[\Gamma][M]k \rightsquigarrow^{\ast} \mathcal{K}[\Gamma][N]k\).
2. If \(\Gamma \vdash C \) and \(C \rightarrow C'\), then there exist well-oriented \(C''\) and \(C'''\) such that \(\mathcal{K}[C''']k \rightarrow^{\ast} \mathcal{K}[C'']k\).

Thus we can simulate concurrent communication using only Functional \(\mu GV\). As a corollary, we obtain that \(\mu GV\) is strongly normalising.

**Theorem 11 (Strong normalisation).** If \(\Gamma \vdash C : T\), then there are
no infinite \(\equiv \rightarrow\) sequences starting from \(C\).

The proof follows immediately from Theorem 10 and the quite
standard result that Functional \(\mu GV\) (linear \(\lambda GV\)-calculus with
positive (co)recursive data types) is strongly normalising. (To show the
latter, map functional \(\mu GV\) into System F, forgetting linearity,
and encoding the positive (co)recursive data types using polymorphism.)
An immediate consequence is that our calculus is free from
livelock; that is, that there are no oscillating sequences of configurations
that diverge.

The strong normalisation result straightforwardly extends to the
extension of \(\mu GV\) with the exponential modality \([26]\), and to the
setting where we allow reduction under lambdas in the source

---

**Covariant Functors**

\[
\begin{align*}
\langle F,G \rangle (M) & = \lambda c.\text{fork}(\lambda d.\text{let } (z, d) = \text{receive } d \text{ in } \text{let } c = \text{send } (F'(M) z, c) \text{ in } \text{link } (G(M) d, G(M) c)) \\
\langle ?F,G \rangle (M) & = \lambda c.\text{fork}(\lambda d.\text{let } (z, c) = \text{receive } c \text{ in } \text{let } d = \text{send } (F(M) z, d) \text{ in } \text{link } (G(M) d, G(M) c)) \\
\langle ?F.\text{end},G \rangle (M) & = \lambda c.\text{end}''(F(M)(\lambda c) c)) \\
\langle ?F.\text{end} \rangle (M) & = \lambda c.\text{fork}(\lambda d.\text{let } (z, d) = \text{receive } d \text{ in } \text{let } c = \text{send } (F(M) z, c) \text{ in } \text{link } (G(M) d, G(M) c))
\end{align*}
\]

**Contravariant Functors**

\[
\begin{align*}
\langle F,G \rangle (\lambda c.\text{fork}(\lambda d.\text{let } (z, d) = \text{receive } d \text{ in } \text{let } c = \text{send } (F'(M) z, c) \text{ in } \text{link } (G(M) d, G(M) c)) \\
\langle ?F,G \rangle (\lambda c.\text{fork}(\lambda d.\text{let } (z, c) = \text{receive } c \text{ in } \text{let } d = \text{send } (F(M) z, d) \text{ in } \text{link } (G(M) d, G(M) c)) \\
\langle ?F.\text{end},G \rangle (\lambda c.\text{end}''(F(M)(\lambda c) c)) \\
\langle ?F.\text{end} \rangle (\lambda c.\text{fork}(\lambda d.\text{let } (z, d) = \text{receive } d \text{ in } \text{let } c = \text{send } (F(M) z, c) \text{ in } \text{link } (G(M) d, G(M) c))
\end{align*}
\]

**Configuration Equivalence**

\[
\begin{align*}
H[\text{link } (x, y)] & \equiv H[\text{link } (y, x)] \\
\text{fork } (\lambda x. M) & \equiv (\lambda z. M) \parallel \lambda y. \text{fork } (\lambda y. M) \\
\text{receive } x & \equiv \lambda y. \text{receive } y \\
\text{send } x & \equiv \lambda y. \text{send } y
\end{align*}
\]

**Configuration Reduction**

\[
\begin{align*}
\text{SEND} & = (\text{new } x)(H[\text{receive } x] \parallel H'[\text{send } (V, x)]) \\
\text{FORK} & = H[\text{fork } (\lambda y. M)] = (\text{new } x)(H[x] \parallel \sigma \text{fork } (\lambda y. M)) \\
\text{WAIT} & = (\text{new } x)(H[\text{wait } x] \parallel \sigma x) \\
\text{LINK0} & = H[\text{link } (x, y)] \equiv (\text{new } x) (\text{link } (x, y) \parallel \text{link } (x, y)) \\
\text{LINK1} & = (\text{new } x, z) (\text{link } (x, y) \parallel \text{link } (x, y)) \parallel \phi M \equiv \phi M \parallel \phi M
\end{align*}
\]

---

**Figure 8:** Concurrent Semantics of \(\mu GV\): Functors, Equivalences, and Reductions.
As well as providing a means to prove termination, the CPS transformation is interesting in its own right as it provides insights into the restricted nature of the concurrency provided by \( \mu \)GV. Furthermore, by composing the CPS translation with the translation from \( \mu \)CP to \( \mu \)GV (§6.3), we obtain a translation from \( \mu \)CP into a typed lambda calculus. This shows that the concurrency of \( \mu \)CP is equivalent to that provided by full reduction in the \( \lambda \)-calculus.

We can make the translation more uniform by factoring it through a polarization phase. In particular, polarization allows us to give a single translation for each of the send, fork, and link cases. Polarization provides a way of encoding output session types as input session types and vice-versa. It also leads to a clean way of handling polymorphic session types by uniformly choosing either a positive (output) or a negative (input) representation for session type variables.

6. The \( \mu \)CP Language

In this section, we present the syntax (§6.1) and semantics (§6.2) of \( \mu \)CP, an extension of Waldner’s \( \mu \)CALCULUS with recursive and corecursive types following Baedel’s approach to recursion and corecursion in classical linear logic [3]. We then argue that \( \mu \)CP and \( \mu \)GV are equally expressive via semantics-preserving translations from \( \mu \)CP into \( \mu \)GV and vice versa.

Figure 9: CPS Translation for Core Types

\[
\begin{align*}
K[T \to U] &= K[T] \to (K[U] \to R) \to R \\
K[T \circ U] &= K[T] \circ K[U] \\
K[T \circ U] &= K[T] \circ K[U] \\
K[T \& U] &= K[T] \& K[U] \\
K[T \& U] &= K[T] \& K[U] \\
K[\mu X.T] &= \mu X. K[T] \\
K[\nu X.T] &= \nu X. K[T] \\
K[X] &= X
\end{align*}
\]

Figure 10: Naive CPS Translation for Core Terms

\[
\begin{align*}
N[\lambda x. M] &= \lambda y. N[M] \\
N[M \times N] &= N[M] \times N[N] \\
N[M \otimes N] &= N[M] \otimes N[N] \\
N[\text{let } (x, y) = M \text{ in } N] &= N[M](\lambda (x, y) \to z) \\
N[\text{case } M \{ \text{inl } x \to N_1; \text{inr } y \to N_2 \}] &= N[M](\lambda x \to N_1; \lambda y \to N_2) \\
N[\text{let } (x, y) = M \text{ in } N] &= N[M](\lambda (x, y) \to z) \\
N[\text{case } M \{ \text{inl } x \to N_1; \text{inr } y \to N_2 \}] &= N[M](\lambda x \to N_1; \lambda y \to N_2)
\end{align*}
\]

Figure 11: CPS Translation for Concurrent \( \mu \)GV and Contexts

\[
\begin{align*}
K[\text{send}_x] &= \text{let } (x, c) = p \text{ in } (c x) k \\
K[\text{send}_y] &= \text{let } (x, c) = p \text{ in } (c x) k \\
K[\text{receive}_c] &= \text{let } (x, c) = p \text{ in } (c x) k
\end{align*}
\]

Shallow Polarization

\[
\begin{align*}
S_1 := \text{T.S} \mid \text{end}_1 \quad S_2 := \text{T.S} \mid \text{end}_2
\end{align*}
\]

\[
\begin{align*}
\text{send}_x : T \to !T.S \mid \text{end}_1 \\
\text{send}_y : T \to !T.S \mid \text{end}_2
\end{align*}
\]

\[
\begin{align*}
\text{fork}_x : (S_1 \to \text{end}_1) \to \text{end}_1 \\
\text{fork}_y : (S_1 \to \text{end}_1) \to \text{end}_1
\end{align*}
\]

Constructions

\[
\begin{align*}
K[\text{let } x = \text{inr } p \text{ in } c] &= K[x] \\
K[\text{let } x = \text{inl } p \text{ in } c] &= K[x]
\end{align*}
\]

Figure 12: CPS Translation of \( \mu \)GV Contexts

6.1 Syntax

Figure 13 gives the terms and typing of \( \mu \)CP, an extension of Waldner’s process calculus CP with recursive and corecursive types. The syntax of types is that of the propositions of linear logic, extended with least (\( \mu F \)) and greatest (\( \nu F \)) fixed points. As in \( \mu \)GV, we have omitted polymorphism and the exponential modality; their reintroduction is entirely orthogonal to our development. The definition of duality includes the duality of least and greatest fixed points; the dual of an operator is defined by \( F^\perp(X) = (F(X^\perp))^\perp \), as for \( \mu \)GV. The terms of \( \mu \)CP are restricted compared to \( \pi \)-calculus in several ways. Most significantly, composition and name restriction are combined in a single syntactic form, and the composed processes are limited to share only the newly introduced name. The forwarding construct \( x \leftrightarrow y \) corresponds to the axiom rule in linear logic; it is necessary for the treatment of recursion and for the extension of \( \mu \)CP to include polymorphism.

A Simpler Send

The \( \mu \)CP rule for output is appealing because it corresponds exactly to the linear logic proof rule for \( \otimes \). Its correspondence to \( \pi \)-calculus is less direct: the term for output includes name restriction (introducing new name \( y \)), output of \( y \) along \( x \), and finally a restricted composition (of \( P \) and \( Q \)). This
Note that the hypotheses of the right rule for \( \mu \) and the left rule for \( \nu \). This leaves us with only two rules, with term assignments as follows:

\[
\begin{align*}
P \vdash y : A, x : F(A) & \quad \text{corec } x(y).P \vdash y : A^+, x : \nu F \\
A & \quad \text{rec } x.P \vdash \Psi : x : \mu F
\end{align*}
\]

However, there is a problem with this formulation. Suppose that we have some term \( Q \vdash x \). We then have the composition new \( x \) (corec \( x(y).P \)) \( x : \nu F \). However, we have no hope of reducing this cut, as we have no rule which can prove \( \nu F \) in isolation. We can address this problem by suspending the cut in question, moving it into the \( \nu \) rule and giving the rule in Figure 13. In our prior work [26] we observe a similar pattern in comparing the \( \otimes \) rule to the typical process calculus rule for output. As in that case, the version without the suspended cut may expose reductions not present in the suspended version. Nevertheless, we can still define the simpler term as syntactic sugar:

\[
\text{corec } x(y).P = \text{corec } x(y)(y \leftrightarrow z \mid P)
\]

**Examples.** We return to the example of natural numbers to give some flavor of the use of recursion and corecursion in \( \mu CP \). We can define the type of natural numbers much as before

\[
N(X) = 1 + X
\]

and we can give very similar definitions of the constructors

\[
\begin{align*}
\text{zero} & = \text{rec } x.\text{inl}(x), 0 \\
\text{succ} & = \text{rec } x.\text{inr}(x) \rightarrow y
\end{align*}
\]

with the expected typings \( \text{zero} \vdash x : \text{Nat} \) and \( \text{succ} \vdash x : \text{Nat} \), \( y : \text{Nat}^+ \). We can define the addition operation as follows:

\[
\text{plus} \vdash x : \text{Nat} \rightarrow \text{Nat}, y : \text{Nat} \rightarrow \text{Nat}^+ \quad \text{with the expected typings}
\]

\[
\begin{align*}
\text{plus} \vdash x : \text{Nat} \rightarrow \text{Nat}, y : \text{Nat} \rightarrow \text{Nat}^+ \\
\text{plus} \vdash x : \text{Nat} \rightarrow \text{Nat}, y : \text{Nat} \rightarrow \text{Nat}^+
\end{align*}
\]

will reduce to \( A \).

### 6.2 Semantics

The semantics of \( \mu CP \) are given by the cut reduction rules in classical linear logic, extended to account for recursion and corecursion, as shown in Figure 14. We write \( f_\nu(P) \) for the free names of process \( P \). Terms are identified up to congruence \( \equiv \). Many of the principle cut reductions (\( \rightarrow_{cc} \)) correspond to process calculus reductions. The reduction of input against output is complicated by the implicit name restriction and composition inherent in the term structure for output. The new rule for \( \mu CP \) is for rec against corec and amounts to one unfolding of the corec term. In defining the unfolding, we rely on functoriality for the operators; if \( P \vdash x : A \rightarrow y : B \), then \( \text{map}_y.f(P) \vdash x : F^+(A^+), y : F(B) \). (We show functoriality for the positive combinators; the remaining cases can be obtained by switching the channels in the given cases.) We write \( \rightarrow \) for \( \rightarrow_{cc} \).

**Theorem 12** (Cut elimination). If \( P \vdash \Psi \), then there is some \( P' \) such that \( P \rightarrow P' \) and \( P' \) is not of the form new \( x \) (corec \( x(Q) \)) for any \( x, Q, Q' \).

This result corresponds to the termination and deadlock freedom results for \( \mu GV \): any well-typed process reduces to one that is blocked on external communication. The commuting conversions (\( \rightarrow_{cc} \)) do not correspond to computational steps (and thus, do not
Theorem 16. and 6.3 Translations between We omit the details of the translations here. This is not only for If P
calculi, these translations preserve both typing and semantics. µ
Structural Congruence
\[ x \leftrightarrow y \equiv y \leftrightarrow x \]
\[ \text{new } x \left( P \mid \emptyset \right) \equiv \text{new } x \left( Q \mid P \right) \]
\[ \text{new } y \left( P \mid \text{new } x \left( Q \mid R \right) \right) \equiv \text{new } y \left( \text{new } x \left( P \mid \text{new } y \left( Q \mid R \right) \right) \mid \text{new } x \left( P \mid Q \right) \right) \]
\[ \text{if } P \equiv \text{if } Q \]

Functionality (positive cases)
\[ \text{map}^{x,A}_{\mu,\text{GV}}(x) = x \leftrightarrow y, \quad X \notin \text{FTV}(A) \]
\[ \text{map}^{x,A}_{\mu,\text{CP}}(P) = P \]
\[ \text{map}^{x,A}_{\mu,\text{GV}}(P) = x(x') \cdot y[y'] \cdot \left( \text{map}^y_{x,y'} \left( P \mid x', x', y' \right) \right) \cdot \text{map}^{y}_{x,y}(P) \]

Primary Cut Reductions
\[ \text{new } x \left( x \leftrightarrow y \mid P \right) \rightarrow_{cc} P \mid x \}
\[ \text{new } x \left( x \leftrightarrow y \left| P \mid Q \right| x \left( y \right) \right) \rightarrow_{cc} \text{new } x \left( Q \mid \text{new } y \left( P \mid Q \right) \right) \]
\[ \text{new } x \left( \text{rec } x \cdot R \right) \rightarrow_{cc} \text{new } x \left( P \mid Q \right) \]
\[ \text{new } x \left( \text{corec } x \left( y \right) \left| P \mid Q \right| \right) \rightarrow_{cc} \text{new } x \left( P \mid Q \right) \]

Commuting Conversions
\[ \text{new } z \left( x[y] \left| P \mid Q \right| R \right) \rightarrow_{cc} x \left( y[z] \right) \left| \text{new } z \left( P \mid Q \right) \right| \left( Q \mid R \right) \]
\[ \text{new } z \left( x \left( y \right) \left| P \mid Q \right| R \right) \rightarrow_{cc} \text{new } z \left( P \mid Q \right) \left( Q \mid R \right) \]
\[ \text{new } z \left( x \left| \text{new } z \left( P \mid Q \right) \right| \right) \rightarrow_{cc} \text{new } z \left( P \mid Q \right) \left( Q \mid R \right) \]
\[ \text{new } z \left( x \left| \text{new } z \left( P \mid Q \right) \right| \right) \rightarrow_{cc} \text{new } z \left( P \mid Q \right) \left( Q \mid R \right) \]

Figure 14: \( \mu \text{CP} \) Reduction Rules

correspond to reductions in process calculi), but play a crucial role in cut elimination by moving remaining internal communication behind any external communication.

6.3 Translations between \( \mu \text{CP} \) and \( \mu \text{GV} \)
We conclude our discussion of \( \mu \text{CP} \) by discussing its relationship with \( \mu \text{GV} \). Our previous work [26] considers a functional calculus GV and a process calculus CP, similar to \( \mu \text{GV} \) and \( \mu \text{CP} \) but lacking recursion and corecursion. In that setting we give translations \( \text{GV}[\_ \rightarrow \_] \) and \( \text{CP}[\_ \rightarrow \_] \) from GV configurations to CP terms and vice versa and show that these translations preserve both typing and semantics.

We have devised similar translations between \( \mu \text{GV} \) and \( \mu \text{CP} \). We omit the details of the translations here. This is not only for reasons of space; the details of the translations are identical but for the treatment of recursion, and the recursive forms in \( \mu \text{GV} \) and \( \mu \text{CP} \) are already quite similar. As in the case of the non-recursive calculus, these translations preserve both typing and semantics.

Theorem 13. If \( P \vdash \Psi \), then \( \text{GV}[\Psi] \vdash \text{CP}[P] : \text{end} \).

Theorem 14. If \( P \vdash \Psi \) and \( P \rightarrow_{cc} Q \) then \( \text{GV}[\Psi] \vdash \text{CP}[P] \rightarrow_{+} \equiv_{+} \text{GV}[\Psi] \).

Theorem 15. If \( \Gamma \vdash C : T \), then \( C[C][z] \equiv \Gamma[z] \rightarrow_{+} \Gamma[z] \rightarrow_{+} \).

Theorem 16. If \( \Gamma \vdash C : T \) and \( C \rightarrow_{cc} C' \), then \( C[C][z] \equiv \Gamma[z] \rightarrow_{+} \Gamma[z] \rightarrow_{+} \).

The simulation of \( \mu \text{GV} \) evaluation by \( \mu \text{CP} \) evaluation is straightforward. The simulation of \( \mu \text{CP} \) evaluation by \( \mu \text{GV} \) evaluation relies on the same administrative weak bisimulation results needed to show that the core \( \mu \text{GV} \) calculus simulates its extension with recursive session types, but introduces no other difficulties.

7. Related Work
Session Types and Linear Logic. Session types were originally introduced by Honda [22] as a typing discipline for a CCS-like process calculus. Takeuchi et al. [30] and Honda et al. [23] extended the original approach to include delegation and recursion. Honda’s system relied on a substructural type system, and borrowed some syntax from linear logic, but did not draw a direct connection between the systems nor suggest the connection between the input and output session types and the \( \otimes \) and \( \otimes \) connectives. Abramsky [1] and Bellin and Scott [5] give interpretations of linear logic proofs as \( \pi \)-calculus processes, and of cut elimination as \( \pi \)-calculus reduction. Their interpretations of \( \otimes \) and \( \otimes \) are very different from the interpretations of input and output in session types. Caires and Pfenning [13] give the formal correspondence between session types and linear logic, interpreting the propositions of intuitionistic linear logic as session types, and showing that \( \pi \)-calculus reduction corresponds with cut reduction. As a consequence of the latter correspondence, they show that cut elimination in linear logic proves deadlock freedom for session-typed \( \pi \) calculus terms. Vasconcelos et al. [33] and Gay and Vasconcelos [20] consider functional languages extended with session-typed concurrency. The functional fragments of their calculi are generally less fully featured than ours (for example, they omit sums) while their concurrent fragments include non-determinism and deadlock. Walder [34] presents a process calculus, called CP, similar to that of Caires and Pfenning, but based on classical rather than intuitionistic linear logic. He also gives a functional calculus, called GV, and similar to that of Gay and Vasconcelos. He shows a type-preserving translation from GV to CP; however, his GV is less expressive than CP. Our prior work [26] introduces a more expressive variant of GV, based on Walder’s, and gives both a direct semantics and semantics-preserving translations to and from Walder’s CP.

Recursive and Corecursive Definition. The interpretation of recursive data types, and their connection to recursive functions, has been studied extensively; we highlight the direct precursors of our approach. Goguen et al. [21] proposed initial algebras and their corresponding folds as a means for understanding recursive data types and their use. Meijer et al. [29] characterized the use of both folds and unfolds, among other patterns, in the definition of recursive functional programs. The coincidence of least and greatest fixed points for data type constructors in many models was first observed by Freyd [19]; he argues that this observation justifies the use of such fixed points for recursive data types. Baelde and Miller [4]
first described an extension of linear logic with induction and coinduction, encoded using the exponential modality and second-order quantification. Baelde [3] treats induction and coinduction without encoding; in particular, he gives a cut reduction rule for recursive and corecursive terms, and shows cut elimination directly.

**Recursive Session Types.** There have been several recent developments of recursive session types and their relationship with linear logic. We highlight three closely related to our development.

Toninho et al. [32] present a system with recursive session types based on intuitionistic linear logic extended with corecursion. They arrive at a similar (albeit intuitionistic) typing discipline for corecursive session types to ours (§2.3), and give a direct proof of termination for the resulting system (without encoding). However, their approach differs from ours in several significant ways. First, they treat recursive processes as primitive, and so do not expose the connection with recursive data types. In contrast, we believe that the parallels with data types (and thus, our ability to present a simple core calculus) is one of the principal benefits of our approach. One consequence is that they have only corecursive processes ($\nu^r, \nu^c$ in our notation), but not recursive processes ($\mu^r, \mu^c$) nor the possibility of identifying greatest and least fixed points. Finally, our session types are classical, while theirs are intuitionistic. One consequence of our approach is that we are explicit about the role of duality, and thus identify a new notion of duality for recursive session types, of our approach is that we are explicit about the role of duality, and suggest a syntactic instance of their characterization. A particular concern of their work, absent from ours, is subtyping: a process may offer more choices than those from which its partner selects. However, their definitions are more complex than ours even without considering subtyping; in particular, they rely on partially unfolding recursive types at each computation of their duals.

Bono and Padovani [10] and Bernardi and Hennessy [6] independently observed that the standard definition of duality for recursive session types fails when recursion occurs in a carried type. Bono et al. [11] present a session-type grounded functional language with self-dual recursion. They do not attempt to enforce termination or distinguish recursion and corecursion, but do present a definition of duality similar to the one that we propose for self-dual recursive session types (§3.2). Bernardi et al. [8] systematically study several duality relations, and propose a notion of session typing independent of the particular duality relation. They also give a coinductive characterization of duality, and suggest a syntactic instance of their characterization. A particular concern of their work, absent from ours, is subtyping: a process may offer more choices than those from which its partner selects. However, their definitions are more complex than ours even without considering subtyping; in particular, they rely on partially unfolding recursive types in each computation of their duals.

**8. Future Work**

We have presented a core concurrent linear $\lambda$-calculus with recursive and corecursive data types, and shown how to encode recursive and corecursive session types and processes in this calculus. We have shown that our type system guarantees termination and lock freedom, giving modular proofs which can easily be extended to encompass additional features. We have given a natural, semantically justified approach to extending our system to non-terminating (but still productive) computation. Finally, we have related our calculus to a process calculus based on classical linear logic with induction and coinduction, giving strong logical foundations to our work.

The model of concurrency in $\mu$GV (like much of the work on logically founded session types) is somewhat limited. By ruling out deadlock, livelock, and in particular data races, it also rules out interesting forms of concurrency. For instance, there is no way of modeling a non-deterministic stateful service such as an online book store in which it should be possible for one customer to observe that a different customer bought the last copy of a book. This raises the question: can the logically founded approach be extended to encompass more realistic forms of concurrency? Recently, Atkey et al. [2] demonstrated that conflating dual propositions in a process calculus based on classical linear logic captures non-determinism, shared state, and more expressive concurrency patterns. On the other hand we would like to transfer these results to GV. On the other hand we would like to transfer ideas for the current paper to extend the work of Atkey et al. to incorporate conflated fixed points.

While the development of $\mu$GV in this paper is largely theoretical, we believe it can also inform practical implementations of session types, including our implementation of session types for the Links web programming language [31]. Indeed, the Links implementation of session types is based on a core calculus $F$ (System $F$ with Session Types), a polymorphic variant of GV [27]. We would like to investigate the relationship between $\mu$GV with FST, while also taking into account practical considerations.

In this paper we have focused on binary session types. Carbone et al. [14] give a logical account of multiparty session types based on CP. It would be interesting to adapt this work to the $\mu$GV setting by following our general approach to relating variants of CP and GV.

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**References**


