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Towards native higher-order remote procedure calls

Olle Fredriksson  Dan R. Ghica  Bertram Wheen
University of Birmingham, UK

Abstract
We present a new abstract machine, called DCESH, which models the execution of higher-order programs running in distributed architectures. DCESH implements a native remote higher-order function call across node boundaries. It is a modernised version of SECD enriched with specialised communication features required for implementing the remote procedure call mechanism. The key correctness result is that the termination behaviour of the remote procedure call is indistinguishable (bisimilar) to that of a local call. The correctness proofs and the requisite definitions for DCESH and other related abstract machines are formalised using Agda. We also formalise a generic transactional mechanism for transparently handling failure in DCESHS.

We use the DCESH as a target architecture for compiling a conventional call-by-value functional language ("Floskel") which can be annotated with node information. Conventional benchmarks show that the single-node performance of Floskel is comparable to that of OCaml, a semantically similar language, and that distribution overheads are not excessive.

1. Native RPC and transparent distribution
The remote procedure call (RPC) [7] is a widely used mechanism for higher-level inter-process communication. An RPC is what its name suggests: a call to a procedure located on a remote node. The idea is that an RPC looks and acts like an ordinary procedure call, and performs the necessary message passing under the hood. However, the RPC mechanism tends to be bolted on top of a pre-existing language, as a library for example, rather than being seamlessly integrated into it. This leads to significant syntactic differences between calling a local library function and a remote function. Even if these syntactic differences can be smoothed using stubs that wrap remote calls into local calls [6] important differences still persist, of which the most important is that arguments must be of ground types.

Generalising RPC to all types and incorporating it seamlessly in the language has been considered but dismissed on grounds of potential inefficiencies [40]. However, considering that a number of technologies that trade efficiency for convenience have been rejected on similar grounds (from machine independent languages to functional programming, garbage collection or automated program verification – to name only a few), we decided it is time to revisit this issue. Because the execution of applications is increasingly and rapidly moving from single devices to networks of (often heterogeneous) devices a case can be made that such revisiting is timely.

A native RPC system, offering as an immediate benefit transparent and automated distribution, can be useful in domains where programmer effectiveness is more important than machine efficiency.

We attack this problem in a principled manner. We want the seamless integration of the RPC in the language to be not merely syntactic but semantic as well: the RPC is a native call of our language, on the same level as a conventional local call. This is realised by introducing a new kind of abstract machine which extends the conventional SECD machine [26], or rather a modernised version of it, with communication primitives. These are not general-purpose low-level communication primitives but are especially designed to support the implementation of RPCs. Even though we aim for simplicity first, some of the technicalities, especially the proofs of correctness, are quite intricate. For this reason the abstract machine network framework and its correctness properties are fully formalised using the Agda language [34]. The technical challenge is not just one of handling complicated formalisms but also mathematical, the key correctness proof requiring the adaptation of the step-indexed relations technique [2] to bisimulation.

Note that we are language rather than system oriented. We assume the existence of a run-time infrastructure to handle system-level aspects associated with distribution such as failure detection, load balancing, global reset and initialisation, and so on. From a systems perspective we only deal with failure handling, which is particularly important in a distributed setting, using a general transactional approach.

The networks of abstract machines can serve as a target for the compilation of a conventional call-by-value language which we call Floskel. The interesting new thing about it is that there is almost nothing new about it. It uses a Haskell-like syntax but it has the semantics of the pure fragment of OCaml. Remote procedure calls are indistinguishable from native calls except for an annotation, which can be applied to any sub-term, indicating that it is to be executed on a different node. From the point of view of the language, and of the programmer, this annotation has no syntactic, semantic or typing implication. It is a mere pragma-like directive.

Prior work on native RPCs and seamless distribution is specialised to web programming [11], thus domain-specific. We aim to be as generic as possible in this context. Some existing work uses as a starting point interaction-based semantic paradigms which lend themselves naturally to a communication-centric implementation: Geometry of Interaction [17] and Game Semantics [18]. Such approaches have two significant disadvantages. The exotic operational behaviour makes it impossible to apply known optimisation techniques, and to interact with code compiled conventionally. This can be seen in the low performance of single-node execution of programs compiled using such techniques. On the other hand, the single-node compilation of Floskel is very similar to the conventional compilation of a language such as OCaml, and the bench-

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marks indicate that the overhead required by the RPC run-time is not excessive.

In a nutshell, we believe that this paper can help make the case that functional languages with native RPCs could be a lot more useful than presumed.

1.1 Technical outline and contributions

Our main contributions are in the following areas:

Compiler and run-time We describe the syntax (Sec. 2.1) and implementation (Sec. 2.2) of Floskel, a general-purpose functional language with native RPCs. Our basis is a conventional compiler for such a language, and we show how it is modified to support RPCs and, additionally, ubiquitous functions, i.e. functions available on all nodes. Our benchmarks suggest that Floskel’s performance is comparable to the state of the art OCaml compiler (Sec. 2.4) for single-node execution.

Abstract machines The semantics of a core of Floskel has been formalised in Agda (Sec. 3) in the form of an abstract machine that can be used to guide an implementation. To achieve this we make gradual refinements to a machine, based on Landin’s SECD machine [26], that we call the CESH machine (Sec. 3.1). First we add heaps for dynamically allocating closures, forming the CESH machine (Sec. 3.2); we show that the CES and CESH execution is bisimilar. We then add communication primitives (synchronous and asynchronous) by defining a general form of networks of nodes that run an instance of an underlying abstract machine (Sec. 3.3). Using these networks, we illustrate the idea of subsuming function calls by communication protocols by constructing a degenerate distributed machine, DCESH1 (Sec. 3.4), that decomposes some machine instructions into message passing, but only runs on one node. Execution on the fully distributed CESH machine called DCESH (Sec. 3.5), is shown to be bisimilar to the CESH machine — our main theoretical result. Finally, we model a general-purpose fault-tolerant environment for DCESH-like machines (Sec. 4) by layering a transactional abstract machine that provides a simple commit-and-rollback mechanism for an abstract machine that may unexpectedly fail.

Formalisation in Agda The theorems that we present in this paper have been proved correct in Agda [34], an interactive proof assistant and programming language based on intuitionistic type theory. The definitions and proofs in this paper are intricate, so carrying them out manually would be error-prone, arduous and perhaps unconvincing. Agda has been a helpful tool in producing these proofs, providing a pleasant interactive environment in which to play with alternative definitions. To eliminate another source of error, we do not present our results in informal mathematics; the code blocks in Sec. 3 come directly from the formalisation. The formalisation is organised as follows, where the arrows denote dependence, the lines with ~ symbols bisimulations, and the parenthesised numerals section numbers:

2. Floskel: a location-aware language

2.1 Syntax

At the core of the Floskel language is a call-by-value functional language with user-definable algebraic data types and pattern matching. Floskel is semantically similar to languages in the ML family, and syntactically similar to Haskell. The main syntactic difference between Floskel and Haskell is that pattern matching clauses are given without the leading function name, to avoid repetition, and that type annotations are given after a single colon, as in the following example:

\[
\begin{align*}
\text{map} : (a \to b) \to [a] \to [b] \\
\text{f} \left[ \right] &= [\] \\
\text{f} \left( x : xs \right) &= f x :: \text{map} f \text{xs}
\end{align*}
\]

Node annotations An ordinary function definition, like map, is a ubiquitous function by default. This means that it is made available on all nodes in the system, and a call to such a function is always done locally – a plain old function call.

On the other hand, a function or sub-term defined with a node annotation, such as

\[
\begin{align*}
\text{query} @ \text{Database} : \text{Query} \to \text{Response} \\
x &= ..., \\
\text{q} &= ... \text{ use q} ...
\end{align*}
\]

can be applied to \text{query} yielding \text{f} \text{query} : X.

Node annotations can also be applied to sub-expressions, as in the following example:

\[
\begin{align*}
\text{sum} \left[ \right] &= 0 \\
\text{f} \left( x : xs \right) &= x + \text{sum} \text{xs} \\
\text{xs} @ \text{A} &= ... \\
\text{ys} @ \text{B} &= ... \\
\text{result} @ \text{C} &= \left( \text{sum} \text{xs} \right) @ \text{A} + \left( \text{sum} \text{ys} \right) @ \text{B}
\end{align*}
\]

Here we want to calculate the sum, on node C, of the elements of two lists located on nodes A and B. If the lists are lengthy, it is better to calculate the sums on A and B, and to then send the final sum to C, since this saves us having to send the full lists over the network.

2.2 Compilation

The Floskel compiler [1] currently targets C using the MPI library for communication, though other targets are possible since we do not make use of any exotic features of our target. Any compatible combination of low-level language and message passing library would work. Most of the compiler’s pipeline is standard for a functional language implementation. It works by applying a series of transformations to the source program until reaching a level low enough to be straightforwardly translated to C. Since the source language has pattern matching, it first compiles the pattern matching to simple case trees [4]. Local definitions are then lifted to the top-level using lambda lifting [25], and lastly the program is closure converted [31] to support partially applied functions.

Up until the lambda lifting, a node annotation is a constructor in the abstract syntax tree of the language’s expressions. The lambda lifter lifts such sub-expressions to the top-level such that...
annotations are afterwards associated with definitions (and not expressions).

The main work specific to Floskel is done in the closure conversion and the run-time system that the compiled programs make use of.

Closures For function applications, the closure converter distinguishes between known functions – those that are on the top-level and have a known arity, and unknown functions – those that are provided e.g. as function arguments.

A known function \( f \) that is either ubiquitous or available on the same node as the definition that is being compiled is compiled to an ordinary function call if there are enough arguments. If there are not, and the function is ubiquitous we have to construct a partial application closure, which contains a pointer to the function and the arguments so far. The compiler maintains the invariant that unknown functions are always in the form of a closure, whose general layout is:

\[
g_{\text{ptr}} \ g_{\text{id}} \ \text{arity} \ \text{payload} \ldots
\]

Since the function may require access to the payload of the closure, \( g_{\text{ptr}} \) is a function of \( \text{arity} \ + \ 1 \) when applying a closure \( c \ell \) as above to arguments \( x_1, \ldots, x_{\text{arity}} \), the call becomes \( g_{\text{ptr}} (c \ell, x_1, \ldots, x_{\text{arity}}) \) meaning that the function has access to the payload through \( c \ell \). To construct the initial closure for a partial application of a function \( f \) of \( \text{arity} \) with \( n \) arguments, we have to conform to this rule, so we construct the closure \( (f_{\text{cl}}, f_{\text{id}}, n, y_1, \ldots, y_{\text{arity}}) \) where \( n = \text{arity} - n \) and \( f' \) is a new ubiquitous top-level function defined as follows:

\[
f' \ c \ell \ x_1 \ldots x_n = \text{case } c \ell \ of
\]

\[
(\ldots y_1 \ldots y_{\text{arity}}) \rightarrow f \ (y_1, \ldots, y_{\text{arity}}, x_1, \ldots, x_n)
\]

A family of \( \text{apply} \) functions handle, in a standard way, applications (of \( f \) arguments) of unknown functions by inspecting the arity stored in the closure to decide whether to construct a new partial application closure with the additional arguments or to apply the function.

The field \( f_{\text{id}} \) is an integer identifier assigned to every function at compile-time and used as a system-wide identifier if the function is ubiquitous, or a node-specific identifier if not. If there are \( k \) ubiquitous functions they are assigned the first \( k \) identifiers, and the nodes of the system may use identifiers greater than \( k \) for their respective located functions. Determining if a function is ubiquitous is thus a simple comparison: \( f_{\text{id}} < k \). Additionally, every node has a table of functions that maps ubiquitous or local located function identifiers to local function pointers, which is used by the deserialiser.

If we have a saturated call to a known remote function, we make a call to the function \( \text{rApply} \), defined in the run-time system (to be described). If we have a non-saturated call to a known remote or located function, we construct the closure \( (f_{\text{cl}}, f_{\text{id}}, \text{arity}, y_1, \ldots, y_{\text{arity}}) \) where \( f' \) is a new ubiquitous top-level function defined as follows:

\[
f' \ c \ell \ x_1 \ldots x_n = \text{case } c \ell \ of
\]

\[
(\ldots y_1 \ldots y_{\text{arity}}) \rightarrow
\]

\[
\text{if } \text{myNode} \equiv f_{\text{id}} \text{ then}
\]

\[
\text{lookup } (f_{\text{id}}) \ (y_1, \ldots, y_{\text{arity}}, x_1, \ldots, x_n);
\]

\[
\text{else}
\]

\[
\text{rApply} \ (\text{node}, f_{\text{id}}, y_1, \ldots, y_{\text{arity}}, x_1, \ldots, x_n)
\]

Here \( \text{myNode} \) is the identifier of the node the code is currently being run at. If it is the same node as the node of \( f' \), we can make an ordinary function call by looking up the function corresponding to \( f_{\text{id}} \) in the function table. Otherwise we call the run-time system function \( \text{rApply} \).

In this way, we construct a closure for located functions that looks just like the closure of an ubiquitous function.

2.3 Run-time

The run-time system defines a family of ubiquitous functions \( \text{rApply} \), that, as we saw above, are used for remote procedure calls and to construct closures for located functions. The function takes a function identifier, a node identifier, and \( \text{arity} \) arguments. It serialises the arguments and sends them together with the function identifier to the given node:

\[
\text{rApply } f_{\text{cl}} \ x_1 \ldots x_{\text{arity}} =
\]

\[
send (f_{\text{id}}, \text{serialise } (x_1), \ldots, \text{serialise } (x_{\text{arity}})) \text{ to } f_{\text{node}};
\]

\[
\text{receive answer from } f_{\text{node}} \rightarrow
\]

When the node \( f_{\text{node}} \) receives this message, it looks the function up in its function table, calls it with the deserialised arguments, and sends back the result:

\[
\text{receive } (f_{\text{id}}, y_1, \ldots, y_{\text{arity}}) \text{ from } f_{\text{remoteNode}} \rightarrow
\]

\[
\text{let result } = \text{lookup } (f_{\text{id}}) \ (\text{serialise } (y_1), \ldots, \text{serialise } (y_{\text{arity}}));
\]

\[
in \text{send result to } f_{\text{remoteNode}}
\]

Serialisation In a remote function call the arguments may be values from arbitrary algebraic data types (like lists and trees), in addition to primitive types and functions.

The serialisation of a primitive type is the identity function, while algebraic data-types require a traversal and flattening of the heap structure. We use tags in the lower bits of a value’s field to differentiate between pointers and non-pointers, which makes this flattening straightforward. The interesting part of serialisation is how to handle closures, both in the case of ubiquitous and located functions.

For closures around ubiquitous functions, we serialise the closure almost as is, but use the function identifiers to resolve the function pointer on the receiving node, as the pointer is not guaranteed to be the same on each node.

To handle located functions, the most straightforward implementation is to use “mobilised” closures that work by exchanging the located function with a ubiquitous function that calls \( \text{rApply} \) to perform the remote procedure call – a sort of lambda-lifting for locations. This is what our implementation currently does. Our formalisation will describe an optimised variant of this scheme, which instead saves the closure on the sending node and sends a pointer to that. The optimised scheme means that we do not unnecessarily send closures containing (potentially large) arguments that are going to end up on the node they originated from anyway. The cost of this optimisation, however, is that it requires us to keep track of heap-allocated pointers across node boundaries using distributed garbage collection. The serialisation currently implemented does not require such garbage collection, but may be slow when dealing with large data.

In detail, to serialise a closure

\[
f_{\text{id}} \ f_{\text{arity}} \ f_{\text{payload}} \ldots
\]

we put a placeholder, \( \text{CL} \), in the place of \( f_{\text{id}} \):

\[
f_{\text{id}} \ f_{\text{arity}} \ f_{\text{payload}} \ldots
\]

where \( \text{payload} \) represents the serialised payload and \( \text{CL} \) is a tag that can be used to identify that this is a closure. To deserialise this on the receiving end, we look up the function pointer associated with \( f_{\text{id}} \) in the ubiquitous function table and substitute that for \( \text{CL} \).

2.4 Performance benchmarks

Single-node Before we measure the performance of the implementation of the native RPC, we analyse how the single-node performance is affected by the distribution overhead even if it is not used — is it feasible for a general-purpose language to be based on the DCESH?
<table>
<thead>
<tr>
<th></th>
<th>trees</th>
<th>nqueens</th>
<th>qsort</th>
<th>primes</th>
<th>tak</th>
<th>fib</th>
</tr>
</thead>
<tbody>
<tr>
<td>Floskel</td>
<td>91.2s</td>
<td>12.2s</td>
<td>9.45s</td>
<td>19.3s</td>
<td>16.5s</td>
<td>10.0s</td>
</tr>
<tr>
<td>ocaml</td>
<td>43.0s</td>
<td>3.10s</td>
<td>3.21s</td>
<td>6.67s</td>
<td>2.85s</td>
<td>1.68s</td>
</tr>
<tr>
<td>relative</td>
<td>2.12s</td>
<td>3.94s</td>
<td>2.94s</td>
<td>2.9</td>
<td>5.77</td>
<td>3.95</td>
</tr>
</tbody>
</table>

**Figure 1.** Single-node performance.

<table>
<thead>
<tr>
<th></th>
<th>trees</th>
<th>nqueens</th>
<th>qsort</th>
<th>primes</th>
<th>tak</th>
<th>fib</th>
</tr>
</thead>
<tbody>
<tr>
<td>µs/remote call</td>
<td>618</td>
<td>382</td>
<td>4.77</td>
<td>13.4</td>
<td>6.94</td>
<td>6.87</td>
</tr>
<tr>
<td>H/remote call</td>
<td>1490</td>
<td>25.8</td>
<td>28.1</td>
<td>27.0</td>
<td>32.0</td>
<td>24.0</td>
</tr>
</tbody>
</table>

**Figure 2.** Distribution overheads.

Fig. 1 shows absolute and relative timings of a number of small benchmarks using integers, lists, trees, recursion, and a small amount of output for printing results. We compare the performance of Floskel programs compiled with our compiler, and equivalent OCaml programs compiled using `ocaml`, a high-performance native-code compiler. Since our compiler targets C, we further compile the generated files to native code using `gcc -O2`. We can see that the running time of programs compiled with our compiler is between two and six times greater than that of those compiled with `ocaml`. These results should be viewed in the light of the fact that our compiler only does a minimal amount of optimisation, whereas a considerable amount of time and effort has been put into `ocaml`.

Moreover, our compiler only produces C code rather than assembly. Compiling to C rather than assembly, especially in the style we use, prevents the C compiler from using whole-program optimisations and is, therefore, a serious source of inefficiencies.

**Distribution overhead** We measure the overhead of our implementation of native remote procedure calls by running the same programs as above, but distributed to between two and nine nodes. The distribution is done by adding node annotations in ways that generate large amounts of communication. We run the benchmarks on a single physical computer with local virtual nodes, which generate large amounts of communication. We run the benchmarks using integers, lists, trees, recursion, and a small amount of output for printing results. We compare the performance of Floskel programs compiled with our compiler, and equivalent OCaml programs compiled using `ocaml`, a high-performance native-code compiler. Since our compiler targets C, we further compile the generated files to native code using `gcc -O2`. We can see that the running time of programs compiled with our compiler is between two and six times greater than that of those compiled with `ocaml`. These results should be viewed in the light of the fact that our compiler only does a minimal amount of optimisation, whereas a considerable amount of time and effort has been put into `ocaml`.

Moreover, our compiler only produces C code rather than assembly. Compiling to C rather than assembly, especially in the style we use, prevents the C compiler from using whole-program optimisations and is, therefore, a serious source of inefficiencies.

It is expected that this benchmark depends largely on the kinds of invocations that are done, since it is more costly to serialise and send a long list or a big closure than an integer. The benchmark hints at this; we appear to get a higher cost for remote calls that are big.

An outlier is the `nqueens` benchmark, which does not do remote invocations with large arguments, but still has a high overhead per call, because it intentionally uses many localised functions.

### 3. Abstract machines and networks

Having introduced the programming language, its compiler and its run-time system we now present the theoretical foundation for the correctness of the compiler. We start with the standard abstract machine model of CBV computation, which we refine, in several steps, into increasingly expressive abstract machines with heap and networking capabilities, while showing along the way that correctness is preserved, via bisimulation results. All definitions and theorems are formalised using the proof assistant Agda, the syntax of which we will follow. We give some of the key definitions and examples in Agda syntax, but most of the description of the formalisation is intended as a high-level guide to the Agda code, which is available online [1]. In order to help the reader navigate the code when a significant theorem or lemma is mentioned the fully qualified Agda name is given in a footnote. Note that we shall not formalise the whole of Floskel but only a core language which coincides with Plotkin’s (untyped) call-by-value PCF [37].

#### 3.1 The CES machine

The starting point is a variation of Landin’s SECD machine [26] called Modern SECD [27], which can be traced to the SECD machine variation of Henderson [22] and to the CEK machine of Fellegi [15], which we call CES (Agda module CES). Just like the machine of Henderson, it uses bytecode for the control component of the machine, and just like the CEK it places the continuations that originally resided in the dump directly on the stack, simplifying the machine configurations.

A CES configuration (Config) is a tuple consisting of a fragment of code (Code), an environment (Env), and a stack (Stack). Evaluation begins with an empty stack and environment, and then follows a stack discipline. Sub-terms push their result on the stack so that their super-terms can consume them. When (and if) the evaluation terminates, the program’s result is the sole stack element.

**Source language** The source language has constructors for lambda abstractions (λ t), applications ( t t’), and variables (var n), represented using De Bruijn indices [13], so a variable is a natural number. Additionally, we have natural number literals (lit n), binary operations (op f t t’), and a conditional (if0 t then t’ else t1). Because the language is untyped, we can express a fixed-point combinator without adding additional constructors.

The machine operates on bytecode and does not directly interpret the source terms, so the terms need to be compiled before they can be executed. The main work of compilation is done by the function `compile`, which takes a term t and a fragment of code e used as a postlude. The bold upper-case names (CLOS, VAR, and so on) are the bytecode instructions, which are sequenced using `;`. Instructions can be seen to correspond to the constructs of the source language, sequentialised.

```
compile’ : Term → Code → Code
compile’ (λ t) c = CLOS (compile’ t RET) ; c
compile’ (t $ t’) c = compile’ t (compile’ t’ (APPL ; c))
compile’ (var x) c = VAR x ; c
compile’ (lit n) c = LIT n ; c
compile’ (op f t t’) c = compile’ t’ (OP f ; c)
compile’ (if0 b then t else f) c =
  compile’ b (COND (compile’ t c) (compile’ f c))
```

**Example 3.1.** To compile a term t we supply END as a postlude: `compile t = compile’ t END`. The term `t = (λx. x) λx. x` is compiled as follows:

```
compile ((λ var 0) $ (λ (λ var 1))) = CLOS (VAR 0 ; RET) ;
CLOS (CLOS (VAR 1 ; RET) ; RET) ; APPL ; END
```

Environments (Env) are lists of values (List Value), which are either natural numbers (nat n) or closures (clos el c). A Closure (Closure) is a fragment of code paired with an environment (Code × Env). Stacks (Stack) are lists of stack elements (List StackElem), which are either values (val v) or continuations (cont cl), represented by closures.

Fig. 3 shows the definition of the transition relation for configurations of the CES machine. A note on Agda syntax: The instruction constructor names are overloaded as constructors for the relation; their usage is disambiguated by context. Arguments in
curly braces are *implicit* and can be automatically inferred. Equality (propositional) is written \( \equiv \).

Stack discipline is clear in the definition of the transition relation. When e.g. \( \text{VAR} \) is executed, the CES machine looks up the value of the variable in the environment and pushes it on the stack. A somewhat subtle part of the relation is the interplay between the \( \text{APPL} \) instruction and the \( \text{RET} \) instruction. When performing an application, two values are required on the stack, one of which has to be a closure. The machine enters the closure, adding the value to the environment, and pushes a return continuation on the stack. The code inside a closure will be terminated by a \( \text{RET} \) instruction, so once the machine has finished executing the closure (and thus produced a value on the stack), that value is returned to the continuation.

Note that it will be useful to establish that the CES machine is deterministic.\(^1\)

**Example 3.2.** We trace the execution of Ex. 3.1 defined above, which exemplifies how returning from an application works. Here we write \( \downarrow \text{CES} \) meaning that the machine uses rule \( s \) to transition from \( a \) to \( b \).

```
let \( c_1 = \text{VAR } 0 \); \text{RET} \)
\( c_2 = \text{CLOS } \langle \text{VAR } 1 ; \text{RET} \rangle ; \text{RET} \)
\( c_1 = \text{VAL } \langle \text{CLOS } \langle c_1, [] \rangle \rangle ; c_2 = \text{VAL } \langle \text{CLOS } \langle c_2, [] \rangle \rangle \)
\( \downarrow \text{CES} \) \( \langle \text{CLOS} \rangle \)
\( \downarrow \text{CES} \) \( \langle \text{APPL}; \text{END} \rangle, \langle \text{[c_1]c_2, cl_1} \rangle \)
\( \downarrow \text{CES} \) \( \langle \text{VAR refl}; \text{RET}, \langle cl_2 \rangle, \langle \text{cont } \langle \text{END}, \langle [\] \rangle \rangle \rangle \rangle \)
\( \downarrow \text{CES} \) \( \langle \text{RET} \rangle, \langle \text{[\[\] c_2]} \rangle \)
```

The final result is therefore the second closure, \( c_{l2} \).

The CES machine terminates with a value \( v \), written \( \text{cfg} \downarrow \text{CES} v \) if it, through the reflexive transitive closure of \( \downarrow \text{CES} \), reaches the end of its code fragment with an empty environment, and \( v \) as its sole stack element. It terminates, written \( \text{cfg} \downarrow \text{CES} \) if there exists a value \( v \) such that it terminates with the value \( v \). It diverges, written \( \text{cfg} \uparrow \text{CES} \) if it is possible to take another step from any configuration reachable from the reflexive transitive closure of \( \downarrow \text{CES} \).

We do not prove formally that the compilation of CBV-PCF to the CES machine is correct, as it is a standard result [12].

### 3.2 CESH: A heap machine

In a compiler implementation of the CES machine targeting a low-level language, closures have to be dynamically allocated in a heap. However, the CES machine does not make this dynamic allocation explicit. Now we make it explicit in a new machine, called the CESH, which is a CES machine with an extra heap component in its configuration.

While heaps are not strictly necessary for a *presentation* of the CES machine, they are of great importance to us. The distributed machine that we will later define needs heaps for persistent storage of data, and the CESH machine forms an intermediate step between that and the CES machine. A CESH configuration is defined as

\[ \text{Config} = \text{Code} \times \text{Env} \times \text{Stack} \times \text{Heap} \]

where Heap is a type constructor for heaps parameterised by the type of its content. The only difference in the definition of the configuration constituents, compared to the CES machine, is that a closure value (the clos constructor of the Value type) does not contain an actual closure, but just a pointer (\( \text{Ptr} \)). The stack is as in the CES machine.

Fig. 4 shows those rules of the CESH machine that are significantly different from the CES: \( \text{CLOS} \) and \( \text{APPL} \). To build a closure, the CESH allocates it in the heap, using the \( \text{new} \) function, which, given a heap and an element, gives back an updated heap and a pointer to the element. When performing an application, the machine has a pointer to a closure, so it looks it up in the heap using the \( \text{get} \) function, which, given a heap and a pointer, gives back the element that the pointer points to (if it exists).

A CESH configuration \( \text{cfg} \) can terminate with a value \( v \), written as \( \text{cfg} \downarrow \text{CESH} v \), terminate (\( \text{cfg} \uparrow \text{CESH} \)), or diverge (\( \text{cfg} \uparrow \text{CESH} \)). These are analogous to the definitions for the CES machine, except that the CESH machine is allowed to terminate with *any* heap.

### 3.2.1 Correctness

To show that our definition of the machine is correct, we construct a bisimulation between the CES and CESH, which given the similarity between the two machines, is almost equality. The difference is dealing with closure values, since the CESH stores pointers rather than closures. The relation for closure values must be parametrised by the heap of the CESH configuration, where the (dereferenced) value of the closure pointer is related to the CES closure.

Formally, the relation is constructed separately for the different components of the machine configurations. For bytecode it is equality, and for closures it is defined component-wise. Values are related only if they have the same heap constructor and related constituents: if the two values are number literals, they are related if they are equal; a CES closure and a pointer are related only if the pointer leads to a CESH closure that is in turn related to the CES closure. Environments are related if they have the same list spine and their values are pointwise related. The relation on stacks is defined similarly, using the relation on values and continuations. Finally, two configurations are \( R_{\text{CES}} \)-related if their components are related.

In the formalisation we define heaps and their properties *abstractly*, rather than using a specific heap implementation. The first key property we require is that dereferencing a pointer in a heap

\[^1\] CES.Properties.determinism-CES
where that pointer was just allocated with a value gives back the same value:
\[ \forall h \ x \to \text{let} (h', \ \text{ptr} \to) = h \to x \in h'! \ \text{ptr} \to \text{just} x \]

We will require a preorder \(\subseteq\) for sub-heaps. The intuitive reading for \(h \subseteq h'\) is that \(h'\) can be used where \(h\) can, i.e. that \(h'\) contains at least the allocations of \(h\). The second key property that we require of a heap implementation is that allocation does not overwrite any previously allocated memory cells (\(\text{proj}\) means first projection):
\[ \forall h \ x \to h \subseteq \text{proj}_1 (h \to x) \]

For any two heaps \(h\) and \(h'\) such that \(h \subseteq h'\), if \(R_{\text{CESH}} \text{cfg} (c, e, s, h)\), then \(R_{\text{CESH}} \text{cfg} (c, e, s, h')\).\(^2\)

Our first correctness result is that \(R_{\text{CESH}}\) is a simulation relation.\(^3\)

The proof is by cases on the sets \(\text{Node}_{\text{soup}}\), representing messages currently in transit, that nodes can time. Asynchronous communication is modelled using a "message to be ready to send and receive a message at a single point in time."

Rendezvous is modelled by \(\tau\)-transition. The property \(\sigma\) is a bisimulation, by which decomposes all application and return instructions into communication. In the example, the function \(\text{APPL}\)-receive updates an element of an indexed family (here relying on the decidability of equality – in MPI, they would correspond to the so-called "node ranks"). The type \(\text{Machine}\) is the type of the nodes’ configurations, and \(\text{Msg}\) the type of messages that the machines can send. The presence of the \(\text{Node}\) argument means that the configuration of a node may know about and depend on its own identifier. The type constructor \(\text{Tagged}\) is used to separate different kinds of local transitions: A Tagged Msg can be \(\tau\) (i.e. a silent transition), \(\text{send} \ \text{msg}\), or \(\text{receive} \ \text{msg}\) for \(\text{msg} : \text{Msg}\).

A synchronous network (\(\text{SyncNetwork}\)) is an indexed family of machines, \(\text{Node} \to \text{Machine}\), representing the nodes of the system. An asynchronous network (\(\text{AsyncNetwork}\)) is an indexed family of machines together with a list of pending messages (\(\text{Node} \to \text{Machine} \times \text{List Msg}\)).

Fig. 3 shows the definition of the transition relation for synchronous and asynchronous networks. It uses \(\text{update}\), a function that corresponds to the usual function update (often written \(f \mid x \mapsto y\)) which updates an element of an indexed family (here depending on the decidability of equality of node identifiers).

There are two ways for a synchronous network to make a transition. The first, \(\text{silent-step}\), occurs when a machine in the network makes a transition tagged with \(\tau\), and is allowed at any time. The second, \(\text{comm-step}\), is the aforementioned rendezvous. A node \(s\) first takes a step sending a message, and afterwards a node \(r\) (which can be \(s\)) takes a step receiving the same message. Asynchronous networks only have one rule, \(\text{step}\), which can be used if a node steps with a tagged message that "agrees" with the pending messages. If the node \(\text{receives}\) a message, the message has to be in the list before the transition. If the node \(\text{send}\)s a message, it has to be there after. If the node takes a \(\text{silent}\) step, the list stays the same before and after. This is formalised using a function called \(\text{destag}\), which creates lists of input and output messages from a tagged message.

Asynchronous networks actually subsume synchronous networks.\(^7\) Going in the other direction is not possible in general, but for some specific instances of the underlying transition relation it is, as we will see later.

### 3.4 DCESH\(_1\): A trivially distributed machine

In higher-order distributed programs containing location specifiers, we will sometimes encounter situations where a function is not available locally. For example, when evaluating the function \(f\) in the term \(f (\text{let} g A) (g \times B)\), we may need to apply the remotely available function \(g\). Our general idea is to do this by decomposing some instructions into communication. In the example, the function \(f\) may send a message requesting the evaluation of \(g\), meaning that the \(\text{APPL}\) instruction is split into a pair of instructions: \(\text{APPL-send}\) and \(\text{APPL-receive}\).

This section outlines an abstract machine, called DCESH\(_1\), which decomposes all application and return instructions into communication. The machine is trivially distributed, because it runs as the sole node in a network, sending messages only to itself. Although it is not used as an intermediate step for the proofs, it is included because it illustrates this decomposition.

A configuration of the DCESH\(_1\) machine (\(\text{Machine}\)) is a tuple consisting of a possibly running thread (\(\text{Maybe Thread}\)), a closure heap (\(\text{Heap Closure}\)), and a "continuation heap" (\(\text{Heap Closure} \times \text{Stack}\)). Since the language is sequential we have at most one thread

\(^2\) CESH.Simulation.HeapUpdate.config

\(^3\) CESH.Simulation.presimulation

\(^4\) Relation.presimulation-to-simulation

\(^5\) CESH.Bisimulation.termination-agrees, CESH.Bisimulation.divergence-agrees

\(^7\) Network.\(\to\)Sync-to.\(\to\)Async
running at once. The thread resembles a CES configuration, Thread = Code \times Env \times Stack, but stacks are defined differently. A stack is now a list of values paired with an optional pointer into the continuation heap. Stack = List Val \times Maybe ContPtr (ContPtr is a synonym for Ptr). When performing an application, when CES would push a continuation on the stack, the DCESH machine is going to stop the current thread and send a message, meaning that it has to save the continuation and the remainder of the stack in the heap for them to persist the thread’s lifetime.

The optional pointer in Stack is an element at the bottom of the list of values. Comparing it to the definition of the CES machine, where stacks are lists of either values or continuations (just closures), we can picture their relation: Whereas the CES machine stores the values and continuations in a single, contiguous stack, the DCESH machine stores first a contiguous block of values until reaching a continuation, at which point it stores (just) a pointer to the continuation closure and the rest of the stack.

The definition of closures, values, and environments are otherwise just like in the CESH machine. The machine communicates with itself using two kinds of messages, APPL and RET, corresponding to the instructions that we are replacing with communication.

Fig. 6 defines the transition relation for the DCESH machine, written \( m \xrightarrow{\text{msg}} m' \) for a tagged message msg and machine configurations \( m \) and \( m' \). Most transitions are the same as in the CESH machine, just framed with the additional heaps and the just meaning that the thread is running. We elide them for brevity.

The interesting rules are the decomposed application and return rules. When an application is performed, an APPL message containing a pointer to the closure to apply, the argument value and a pointer to a return continuation (which is first allocated) is sent, and the thread is stopped (nothing). We call such a machine inactive. The machine can receive an application message if the thread is not running. When that happens, the closure pointer is dereferenced and entered, adding the received argument to the environment. The stack is left empty apart from the continuation pointer of the received message. When returning from a function application, the machine sends a return message containing the continuation pointer and the value to return.

On the receiving end of that communication, it dereferences the continuation pointer and enters it, putting the result value on top of the stack.

**Example 3.3.** We trace the execution of Ex. 3.1 in a synchronous network of nodes indexed by the unit type. Heaps with pointer mappings are written \( \{ \text{ptr} \rightarrow \text{element} \} \). The last list shown in each step is the message list of the asynchronous network.

\[
\begin{align*}
\text{let } h_{i} & = \{ \text{ptr}_{1} \rightarrow (c_{1}, [\ ]), \} \\
& \quad \{ \text{ptr}_{2} \rightarrow (c_{2}, [\ ]), \} \\
\text{h}_{0} & = \{ \text{ptr}_{1} \rightarrow ((\text{END}, [\ ]), [\ ]), [\ ] \} \\
& \quad \{ \text{ptr}_{2} \rightarrow (\text{RET}, [\ ]), [\ ] \} \\
\text{in} \quad & (\text{CLOS } c_{1} : \text{CLOS } c_{2} : \text{APPL} : \text{END} : [\ ] : \{ \] : \text{nothing} : h_{i} : 0 : 0 : [\ ]} \\
& \quad \rightarrow (\text{step CLOS}) \\
& (\text{just } (\text{CLOS } c_{2} : \text{APPL} : \text{END} : [\ ] : \{ \text{clos } \text{ptr}_{2} : \text{nothing} : h_{i} : 0 : 0 : [\ ]} \\
& \quad \rightarrow (\text{step CLOS}) \\
& (\text{just } (\text{APPL} ; \text{END} : [\ ] : \{ \text{clos } \text{ptr}_{2} : \text{closer}_{1} : \text{nothing} : h_{i} : 0 : 0 : [\ ]} \\
& \quad \rightarrow (\text{step APPL-send}) \\
& (\text{nothing}, h_{i} : h_{o} : \{ \text{APPL } \text{ptr}_{2} \{ \text{clos } \text{ptr}_{2} \} : \text{nothing} : h_{i} : 0 : 0 : [\ ]} \\
& \quad \rightarrow (\text{step APPL-receive}) \\
& (\text{just } (\text{VAR } \theta : \text{RET} : \{ \text{closer}_{2} : \text{nothing} : h_{i} : 0 : 0 : [\ ]}) \\
& \quad \rightarrow (\text{step VAR-ref}) \\
& (\text{just } (\text{RET} : \{ \text{clos } \text{ptr}_{2} : \text{nothing} : h_{i} : 0 : 0 : [\ ]}) \\
& \quad \rightarrow (\text{step RET-send}) \\
& (\text{nothing}, h_{i} : h_{o} : \{ \text{RET } \text{ptr}_{2} : \text{nothing} : h_{i} : 0 : 0 : [\ ]}) \\
& \quad \rightarrow (\text{step RET-receive}) \\
& (\text{just } (\text{END} : [\ ] : \{ \text{clos } \text{ptr}_{2} : \text{nothing} : h_{i} : 0 : 0 : [\ ]}) \\
& \quad \rightarrow (\text{step CLOS})
\end{align*}
\]

Comparing this to Example 3.2 we can see that an APPL-send followed by an APPL-receive amounts to the same thing as the APPL rule in the CES machine, and similarly for the RET instruction.

### 3.5 DCESH: The distributed CESH machine

We have so far seen two refinements of the CES machine. We have seen CESH, that adds heaps, and DCESH, that decomposes instructions into communication in a degenerate network of only one node. Our final refinement is a distributed machine, DCESH, that supports multiple nodes. The main problem that we now face is that there is no centralised heap, but each node has its own local heap. This means that, for supporting higher-order functions across node boundaries, we have to somehow keep references to closures in the heaps of other nodes. Another problem is efficiency; we would like a system where we do not pay the higher price of communication for locally running code. The main idea for solving these two problems is to use remote pointers, RPtr = Ptx Node, pointers paired with node identifiers signifying what node’s heap the pointer is located. This solves the heap problem because we always know where a pointer comes from. It can also be used to solve the efficiency problem since we can choose what instructions to run based on whether a pointer is local or remote. If it is local, we run the rules of the CESH machine. If it is remote, we run the decomposed rules of the DCESH machine.

The final extension to the term language and bytecode will add support for location specifiers. We add a term construct \( t @ i \), and an instruction REMOTE \( c \ i \) for its compilation. The location specifiers, \( r @ i \), are taken to mean that the term \( r \) should be evaluated on node \( i \). For compilation, we require that the terms \( r \) in all location specification sub-terms \( \{ r @ i \} \) are closed. Terms where this does not hold are transformed automatically using lambda lifting [25] (transform every sub-term \( r @ i \) to \( r = ((\lambda x : t @ i \ i) i) (f v i) \)). The REMOTE \( c \ i \) instruction will be used to start running a code fragment \( c \) on node \( i \) in the network. We also extend the compile' function to handle the new term construct:

\[ \text{compile'} (t @ i) c = \text{REMOTE} (\text{compile'} t \text{ RET} i) \ i \ c \]

Note that we reuse the RET instruction to return from a remote computation.

The definition of closures, values, environments and closure heaps are the same as in the CESH machine, but using RPtr instead of Ptx for closure pointers.
The stack combines the functionality of the CES(H) machine, permitting local continuations, with that of the DCESH₁ machine, making it possible for a stack to end with a continuation on another node. A stack element is a value or a (local) continuation signified by the val and cont constructors. A stack (Stack) is a list of stack elements, possibly ending with a (remote) pointer to a continuation, making it possible for a stack to end with a continuation on another node.

Fig. 7 defines the transition relation of the DCESH₁ machine, written \( i \vdash m \vdash m' \) for a node identifier \( i \), a tagged message \( m \) and machine configurations \( m \) and \( m' \). The parameter \( i \) is taken to be the identifier of the node on which the transition is taking place. For local computations, we have rules analogous to those of the CESH₁ machine, so we omit them and show only those for remote computations. The rules use the function \( i \vdash h \triangleright x \) for allocating a pointer to \( x \) in a heap \( h \) and then constructing a remote pointer tagged with node identifier \( i \) from it. When starting a remote computation, the machine allocates a continuation in the heap and sends a message containing the code and continuation pointer to the remote node in question. Afterwards the current thread is stopped.

On the receiving end of such a communication, a new thread is started, placing the continuation pointer at the bottom of the stack for the later return to the caller node. To run the apply instruction when the function closure is remote, i.e. its location is not equal to the current node, the machine sends a message containing the closure pointer, argument value, and continuation, like in the DCESH₁ machine. On the other end of such a communication, the machine dereferences the pointer and enters the closure with the received value. The bottom remote continuation pointer is set to the received continuation pointer. After either a remote invocation or a remote application, the machine can return if it has produced a value on the stack and has a remote continuation at the bottom of the stack. To do this, a message containing the continuation pointer and the return value is sent to the location of the continuation pointer. When receiving a return message, the continuation pointer is dereferenced and entered with the received value.

A network of abstract machines is obtained by instantiating the Network module with the AsyncNetwork and SyncNetwork and their transition relations refer to the instantiated versions.

Unsurprisingly, if all nodes in a synchronous network except one are inactive, then the next step is deterministic. Another key ancillary property of DCESH networks is that synchronous or asynchronous networks for single threaded computations behave essentially the same, which means it is enough to deal with the simpler synchronous networks.

DCESH network nodes can terminate with a value \( v \) (nodes \( \downarrow \text{S} \) for \( \text{Synchronous} \) terminating nodes), or diverge (nodes \( \uparrow \text{S} \)). A network terminates with a value \( v \) if it can step to a network where only one node is active, and that node has reached the END instruction with the value \( v \) on top of its stack. The other definitions are analogous to those of the CES(H) machine.

### 3.5.1 Correctness

To prove the correctness of the machine, we will now establish a bisimulation between the CESH and the DCESH₁ machines.

To simplify this development, we extend the CESH machine with a dummy rule for the REMOTE \( c \) \( i \) instruction so that both machines run the same bytecode. This rule is almost a no-op, but since we are assuming that the code we run remotely is closed, the environment is emptied, and since the compiled code \( c \) will end in a RET instruction a return continuation is pushed on the stack.

\[
\text{REMOTE } c' \vdash i ; c, e, s, h \quad \xrightarrow{\text{CESH}} \quad (c', \_), \text{cont } (c, e) :: s, h
\]

The relation that we are about to define is, as before, almost equality. But since values may be pointers to closures, it must be parameterised by heaps. A technical problem is that both machines use pointers, and the DCESH machine also uses remote pointers and has two heaps for each node: so the relation must be parameterised by all the heaps in the system. The extra parameter is a synonym for an indexed family of the closure and continuation heaps, \( \text{Heaps} = \text{Node} \rightarrow \text{DCESH.ClosHeap} \times \text{DCESH.ContHeap} \). The complexity of this relation justifies our use of mechanised reasoning.

The correctness proof itself is not routine. Simply following the recipe that we used before does not work. In the old proof, there can be no circularity, since that bisimulation was constructed inductively on the structure of the CES configuration. But now both systems, CESH and DCESH, have heaps where there is a potential
for circular references (e.g. a closure, residing in a heap, whose environment contains a pointer to itself), preventing a direct proof via structural induction. This is perhaps the most mathematically (and formally) challenging point of the paper. The solution lies in using the technique of step-indexed relations, adapted to the context of bisimulation relations [2]. The additional rank parameter records how many times pointers are allowed to be dereferenced.

The rank is used in defining the relation for closure pointers $R_{ptr}$. If the rank is zero, the relation is trivially fulfilled. If the rank is non-zero then three conditions must hold. First, the CESH pointer must point to a closure in the CESH heap; second, the remote pointer of the DCESH network must point to a closure in the heap of the location that the pointer refers to; third, the two closures must be related. The relation for stack elements $R_{StackEl}$. is almost as before, but now requires that the relation is true for any natural number rank, i.e. for any finite number of pointer dereferences. The relation for stacks $R_{Stack}$ now takes into account that the DCESH stacks may end in a pointer representing a remote continuation, requiring that the pointer points to something in the DCESH stacks may end in a pointer representing a remote

Figure 7. The definition of the transition relation of the DCESH machine (excerpt).

4. Fault-tolerance via transactions

In this section we present a generic transaction-based method for handling failure which is suitable for the DCESH. Node state is "backed up" (commit) at certain points in the execution, and if an exceptional condition arises, the backup is restored (roll-back).

This development is independent of the underlying transition relation, but the proofs rely on sequentiality. We assume that we have two arbitrary types $Machine$ and $Msg$, as well as a transition relation over them:

\[ \tau : Machine \rightarrow Tagged Msg \rightarrow Machine \rightarrow * \]

Since we have no knowledge of exceptional states in $Machine$, since it is a parameter, we define another relation, $\tau_{crash} : Machine \rightarrow \omega$ as a thin layer on top of $\tau$. The new definition is shown in Fig. 8 and adds the exceptional state $nothing$ by extending the set of states of the relation to $Maybe Machine$. The fallible machine can make a normal-step transition from and to just ordinary $Machine$ states, or it can crash which leaves it in the exceptional state. This means that we tolerate fail-stop faults as opposed to e.g. the more general Byzantine failures.

The additional assumptions for sequentiality are that we have a decidable predicate, active : $Machine \rightarrow *$ on machines, and the following functions:

\[ inactive-receive-active : \forall \{ m \rightarrow m' \rightarrow \omega \} \rightarrow m ( receive mstg m' ) \rightarrow \omega \]

\[ active-silent-active : \forall \{ m \rightarrow \omega \} \rightarrow m \]

\[ \tau_{crash} : \forall \{ m \rightarrow Crash \} \rightarrow m \]

\[ \tau_{normal-step} : \forall \{ m \rightarrow m' \rightarrow Crash \} \rightarrow \omega \]

\[ \tau_{crash} : \forall \{ m \rightarrow Crash \} \rightarrow m \]

\[ \tau_{sgn} : \forall \{ m \rightarrow Crash \} \rightarrow nothing \]

\[ \tau_{sgn} : \forall \{ m \rightarrow Crash \} \rightarrow nothing \]

Figure 8. The definition of the transition relation of a machine that may crash.
tagged messages that we can observe, namely closure of the above relations. First we define tagged message, as follows:

\[ \text{receive} : \forall \{ m \in m \times \text{msg} \} \rightarrow (m, n) \rightarrow (m, n) \times \text{backup} \]

\[ \text{send} : \forall \{ m \in m \times \text{msg} \} \rightarrow (m, n) \rightarrow (m, n) \times \text{backup} \]

\[ \text{recovery} : \forall \{ m \in m \} \rightarrow (m, n) \rightarrow (m, n) \times \text{backup} \]

These functions express the property that if a machine is invoked, i.e. it receives a message, then it must go from an inactive to an active state. If the machine then takes a silent step, it must remain active, and when it sends a message it must go back to being inactive. This gives us sequentiality; a machine cannot fork new threads, and cannot be invoked several times in parallel.

As the focus here is on obvious correctness and simplicity, we abstract from the method of detecting faults in a separate node, and assume that it can be done (using e.g. a heartbeat network). Similarly, we assume that we have a means of creating and restoring a backup of a node in the system; how this is done depends largely on the underlying system. We so define a machine with a backup as follows:

\[ (m \rightarrow^* m) \rightarrow \text{active} \times \text{active} \]

\[ (m \rightarrow^* m) \rightarrow (m, n) \rightarrow (m, n) \times \text{backup} \]

Using these relations we can define the observable trace of run of a Machine (Backup), i.e. an element of the reflexive transitive closure of the above relations. First we define the subset of tagged messages that we can observe, namely receipt and receive:

\[ \text{data IO (A : *) : * where} \]

\[ \text{send receive : A A IO A} \]

The following function now gives us the observable trace, given an element of the above runtime of a machine (Backup), i.e. an element of the reflexive transitive closure of the above relations. First we define the subset of tagged messages that we can observe, namely receipt and receive:

\[ \text{data IO (A : *) : * where} \]

\[ \text{send receive : A A IO A} \]

The following function now gives us the observable trace, given an element of the above runtime of a machine (Backup), i.e. an element of the reflexive transitive closure of the above relations. First we define the subset of tagged messages that we can observe, namely receipt and receive:

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The following function now gives us the observable trace, given an element of the above runtime of a machine (Backup), i.e. an element of the reflexive transitive closure of the above relations. First we define the subset of tagged messages that we can observe, namely receipt and receive:

\[ \text{data IO (A : *) : * where} \]

\[ \text{send receive : A A IO A} \]
questionable whether acceptable efficiency can be achieved. These arguments certainly apply to our work, where we do just this. With the goal of expressivity in mind, however, we believe that we should enable the programmer to write the potentially inefficient programs that (internally) use remote pointers, because not all programs are performance critical. Furthermore, using a tagged pointer representation [30] for closure pointers means that we can tag pointers that are remote, and pay a very low, if any, performance penalty for local pointers.

Remote Evaluation (REV) [39] is another generalisation of RPC, siding with us on enabling the use of higher-order functions across node boundaries. The main differences between REV and our work is that REV relies on sending code, whereas we can do both, and that it has a more elaborate distribution mechanism.

The well-researched project Eden [29] and the associated abstract machine DREAM [8], which builds on Haskell, is a semi-implicit language. Eden allows expressing distributed algorithms at a high level of abstraction, and is mostly implicit about communication, but explicit about process creation. Eden is specified operationally using a two-level semantics similar to ours, but in the context of the call-by-need evaluation strategy. Kanor [23] is a project that similarly aims to simplify the development of distributed programs by providing a declarative language for specifying communication patterns inside an imperative host language.

Hop [38], Links [11], and ML5 [32] are examples of so called tierless languages that allow writing (for instance) the client and server code of web applications in unified languages with more or less seamless interoperability between them. The Links language shares our goal of unifying distributed programs into a single language with seamless interoperability between the nodes, but its focus is on web programming with client, server, and database. For that purpose it includes sub-languages for elegantly constructing and manipulating XML documents and for doing database queries. The behaviour of Links is specified using the client/server calculus [10], which is an operational semantics similar in purpose to our abstract machines, but, as its name suggests, limited to two nodes. The two nodes of the system are constructed with web-programming in mind and are not equal peers. The server is stateless to be able to handle a large number of clients and unexpected disconnections. The semantics operates on a first-order language and uses explicit substitutions. The client/server calculus also inspired work on the LSAM [33], that lifts the two-node limitation. A similar, but earlier, machine is that of dML [35]. The dML and LSAM machines are conceptually close to our machine, but described on a level that is not as readily implementable as our work, using explicit substitutions and synchronous message passing.

On the language side, our work draws inspiration from abstract machines for game semantics [17, 18] where an exotic compilation technique based on game semantics is used to implement a language like ours, but for call-by-name and without algebraic data types. Recent work shows a formalisation similar to ours, but for call-by-name and without algebraic data types. Recent work shows a formalisation similar to ours, but for call-by-name and without algebraic data types. Recent work shows a formalisation similar to ours, but for call-by-name and without algebraic data types. Recent work shows a formalisation similar to ours, but for call-by-name and without algebraic data types. Recent work shows a formalisation similar to ours, but for call-by-name and without algebraic data types.
References


