Behavioral interface description of an object-oriented language with futures and promises

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ABSTRACT

This paper formalizes the observable interface behavior of a concurrent, object-oriented language with futures and promises. The calculus captures the core of Creol, a language, featuring in particular asynchronous method calls and, since recently, first-class futures.

The focus of the paper are open systems and we formally characterize their behavior in terms of interactions at the interface between the program and its environment. The behavior is given by transitions between typing judgments, where the absent environment is represented abstractly by an assumption context. A particular challenge is the safe treatment of promises: the erroneous situation that a promise is fulfilled twice, i.e., bound to code twice, is prevented by a resource aware type system, enforcing linear use of the write-permission to a promise. We show subject reduction and the soundness of the abstract interface description.

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1. Introduction

How to marry concurrency and object-orientation has been a long-standing issue; (see e.g. [19]) for an early discussion of different design choices. The thread-based model of concurrency, prominently represented by languages like Java and C, has been recently criticized, especially in the context of component-based software development. As the word indicates, components are (software) artifacts intended for composition, i.e., open systems, interacting with a surrounding environment. To compare different concurrency models for open systems on a solid mathematical basis, a semantic description of the interface behavior is needed, and this is what we provide in this work. We present an open semantics for the core of the Creol language [37, 60], an object-oriented, concurrent language, featuring in particular asynchronous method calls and, since recently [39], first-class futures. An open semantics means, it describes the interface behavior of a program or a part of a program, interacting with its environment.

Futures and promises

A future, very generally, represents a result yet to be computed. It acts as a proxy for, or reference to, the delayed result from some piece of code (e.g., a method or a function body in an object-oriented, resp. a functional setting). As the consumer
of the result can proceed its own execution until it actually needs the result, futures provide a natural, lightweight, and (in a functional setting) transparent mechanism to introduce parallelism into a language. Since their introduction in Multilisp [52,21], futures have been used in various languages like Alice ML [64,17,79], E [41], the ASP-calculus [28], Creol, and others. A promise is a generalization\(^1\) insofar as the reference to the result on the one hand, and the code to calculate the result on the other, are not created at the same time; instead, a promise can be created and only later, after possibly passing it around, be bound to the code (the promise is fulfilled).

The notion of futures goes back to functional programming languages. In that setting, futures are annotations to side-effect-free expressions,\(^2\) that can be computed in parallel to the rest of the program. If some program code needs the result of a future, its execution blocks until the future’s evaluation is completed and the result value is automatically fetched back (implicit futures). An important property of future-based functional programs is, that future annotations do not change the functionality: the observable behavior of an annotated program equals the observable behavior of its non-annotated counterpart. This property is no longer assured in the object-oriented setting.

**Interface behavior**

An open program, in this setting, interacts with its environment via message exchange. Besides message passing, of course, different communication and synchronization mechanisms exist (shared variable concurrency, multi-cast, blackboard communication, publish-and-subscribe and many more). We concentrate here, however, on basic message passing using asynchronous method calls. In that setting, the interface behavior of an open program \(C\) can be characterized by the set of all those message sequences (traces) \(t\), for which there exists an environment \(E\) such that \(C\) and \(E\) exchange the messages recorded in \(t\). Thereby we abstract away from any concrete environment, but consider only environments that are compliant to the language restrictions (syntax, type system, etc.). Consequently, interactions are not arbitrary traces \(C \rightarrow_\delta\); instead we consider behaviors \(C \parallel E \Rightarrow \bar{t} \parallel \bar{E}\) where \(E\) is a realizable environment and trace \(\bar{t}\) is complementary to \(t\), i.e., each input is replaced by a matching output and vice versa. The notation \(C \parallel E\) indicates that the component \(C\) runs in parallel with its environment or observer \(E\). To account for the abstract environment (“there exists an \(E\) s.t. . . . “), the open semantics is given in an assumption-commitment way:

\[
\Delta \vdash C : \Theta \quad \Rightarrow \quad \Delta \vdash \hat{C} : \emptyset,
\]

where \(\Delta\) (as an abstract version of \(E\)) contains the assumptions about the environment, and dually \(\Theta\) the commitments of the component. Abstracting away also from \(C\) gives a language characterization by the set of all possible traces between any component and any environment.

Such a behavioral interface description is relevant and useful for the following reasons. (1) The set of possible traces given this way is more restricted (and realistic) than the one obtained when ignoring the environments. When reasoning about the trace-based behavior of a component, e.g., in compositional verification, with a more precise characterization one can carry out stronger arguments. (2) When using the trace description for black-box testing, one can describe test cases in terms of the interface traces and then synthesize appropriate test drivers from it. Clearly, it makes no sense to specify impossible interface behavior, as in this case one cannot generate a corresponding tester. (3) A representation-independent behavior of open programs paves the way for a compositional semantics, a two-level semantics for the nested composition of program components. It allows furthermore optimization of components: only if two components show the same external, observable behavior, one can replace one for the other without changing the interaction with any environment. (4) The formulation gives insight into the semantic nature of the language, here, the externally observable consequences of futures and promises. This helps to compare alternatives, e.g., the Creol concurrency model with Java-like threading.

**Results**

The paper formalizes the abstract interface behavior for concurrent object-oriented languages with futures and promises. The contributions are the following.

**Concurrent object calculus with futures and promises.** We formalize a class-based concurrent language featuring futures and promises. The formalization is given as a typed, imperative object calculus in the style of [11] resp. one of its concurrent extensions. The operational semantics for components distinguishes unobservable component-internal steps from external steps which represent observable component-environment interactions. We present the semantics in a way that facilitates comparison with Java’s multi-threading concurrency model, i.e., the operational semantics is formulated so that the multi-threaded concurrency as (for instance) in Java and the one here based on futures are represented similarly.

---

1 The terminology concerning futures, promises, and related constructs is not too consistent in the literature. Sometimes, the two words are used as synonyms. Interested in the observable differences between futures and promises, we distinguish the concepts and thus follow the terminology as used e.g., in Java, Alice ML, and the definition given in Wikipedia.

2 Though in e.g. Multilisp also expressions with side-effects can be computed in parallel, but still under the restriction that the observable behavior equals that of the sequential counterpart.
Linear type system for promises. The calculus extends the semantic basis of Creol as given for example in [39] with promises. Promises can refer to a computation with code bound to it later, where the binding is done at most once. To guarantee such a write-once policy when passing around promises, we refine the type system introducing two type constructors

\[ T^{+} \quad \text{and} \quad T^{*} \]

The first one represents a reference to a promise that can still be written (and read) with result type \( T \), the second one where only a read-permission is available. The write permission constitutes a resource which is consumed when the promise is fulfilled. The resource-aware type system is therefore formulated in a linear manner wrt. the write permissions. Linear type systems [85] or linear logics [48] are, roughly speaking, instances of so-called sub-structural type systems resp. logics. In constrast to ordinary such derivation systems, where a hypothesis can be used as many times as needed for carrying out a proof, derivation systems built upon linear use of assumptions work differently: using an assumption in a proof step "consumes" it. That feature allows in a natural way to reason about "resources": In our setting, the write-permission is such a resource, and using the corresponding type to derive well-typedness of one part of the program consumes that right such that it is not longer available for type-checking the rest of the program, assuring the intended write-once discipline. The type system resembles in intention the one in [72] for a functional calculus with references. Our work is more general, in that it tackles the problem in an object-oriented setting (which, however, conceptually does not pose much complications), and, more importantly, in that we do not consider closed systems, but open components. Also this aspect of openness is not dealt with in [39]. Additionally, the type system presented here is simpler than in [72], as it avoids the representation of the promise-concept by so-called handled futures.

Soundness of the abstractions. We show soundness of the abstractions, which includes

- **subject reduction**, i.e., preservation of well-typedness under reduction. Subject reduction is not just proven for a closed system (as usual), but for an open system interacting with its environment. Subject reduction implies furthermore:
- **absence of run-time errors** like "message-not-understood", also for open systems.
- **soundness** of the interface behavior characterization, i.e., all possible interaction behavior is included in the abstract interface behavior description.
- for promises: absence of write-errors, i.e. the attempt to fulfill a promise twice.

The paper is organized as follows. Section 2 defines the syntax, the type system, and the operational semantics, split into

### 2.1. Syntax

This section presents the calculus, based on a version of the Creol-language with first-class futures [39] and extended with promises. It is a concurrent variant of an imperative, object-calculus in the style of the calculi from [1]. Our calculus covers first-class futures, which can be seen as a generalization of asynchronous method calls and promises.

We start with the abstract syntax in Section 2.1. After discussing the type system in Section 2.2, we present the operational semantics in Section 2.3.

#### 2.1. Syntax

The abstract syntax is given in Table 1. It distinguishes between user syntax and run-time syntax (the latter underlined). The user syntax contains the phrases in which programs are written; the run-time syntax contains syntactic constituents additionally needed to express the behavior of the executing program in the operational semantics. The latter are not found in a program written by the user, but generated at run-time by the rules of the operational semantics.

The basic syntactic category of names \( n \), which count among the values \( v \), represents references to classes, to objects, and to futures/promises. To facilitate reading, we allow ourselves to write \( o \) and its syntactic variants for names referring to

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<th>Table 1</th>
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<td>Abstract syntax.</td>
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objects, c for classes, and n when being unspecific. Technically, the disambiguation between the different roles of the names is done by the type system and the abstract syntax of Table 1 uses the non-specific n for names. The unit value is represented by () and x stands for variables, i.e., local variables and formal parameters, but not instance variables.

A component C is a collection of classes, objects, and (named) threads, with Ø representing the empty component. The sub-entities of a component are composed using the parallel-construct ||. The entities executing in parallel are the named threads n(t), where t is the code being executed and n the name of the thread. In the given setting, threads are always promises (with the exception of initial threads, see Section 2.3), with their name being the reference under which the result value of t, if any, will be available. A class c(O) carries a name c and defines its methods and fields in O. An object o(c, F, L) with identity o keeps a reference to the class c it instantiates, stores the current value F of its fields, and maintains a binary lock L indicating whether any code is currently active inside the object (in which case the lock is taken) or not (in which case the lock is free). The symbols ⊤, resp., ⊥, indicate that the lock is taken, resp., free. From the three kinds of entities at component level – threads n(t), classes c(O), and objects o(c, F, L) – only the threads are active, executing entities, being the target of the reduction rules. The objects, in contrast, store the state in their fields or instance variables, whereas the classes are constant entities specifying the methods.

The named threads n(t) are incarnations of method bodies “in execution”. Incarnations insofar, as the formal parameters have been replaced by actual ones, especially the method’s self-parameter has been replaced by the identity of the target object of the method call. The term t is basically a sequence of expressions, where the let-construct is used for sequencing and for local declarations. During execution, n(t) contains in t the currently running code of a method body. When evaluated, the thread is of the form n⟨v⟩ and the value can be accessed via n, the future reference, or future for short.

Each thread belongs to one specific object “inside” which it executes, i.e., whose instance variables it has access to. Object locks are used to rule out unprotected concurrent access to the object states: though each object may have more than one method body incarnation partially evaluated, at each time point at most one of those bodies (the lock owner) can be active inside the object. In the terminology of Java, all methods are implicitly considered “synchronized”. The crucial difference between Java’s multi-threading concurrency model and Creol’s active objects model used here is the way method calls are issued at the caller site. In Java and similar languages, method calls are synchronous in the sense that the calling activity blocks to wait for the return of the result and thus the control is transferred to the callee. Here, method calls are issued asynchronously, i.e., the calling thread continues executing and the code of the method being called is computed concurrently in a new thread located in the callee object. In that way, a method call never transfers control from one object, the caller, to another one, the callee. In other words, no thread ever crosses the boundaries of an object, which means, the boundaries of an object are at the same time boundaries of the threads and thus, the objects are at the same time units of concurrency. Thus, the objects are harnessing the activities and can be considered as bearers of the activities. This is typical for object-oriented languages based on active objects.

The final construct at the component level is the v-operator for hiding and dynamic scoping, as known from the π-calculus. In a component C = v(n:T).C’, the scope of the name n (of type T) is restricted to C’ and unknown outside C. v-binders are introduced when dynamically creating new named entities, i.e., when instantiating new objects or new promises. The scope is dynamic, i.e., when the name is communicated by message passing, it is enlarged.

Besides components, the grammar specifies the lower level syntactic constructs, in particular, methods, expressions, and (unnamed) threads, which are basically sequences of expressions. A method ζ(s:T).λ(x:T).t provides the method body t abstracted over the ζ-bound “self” parameter s and the formal parameters x. For uniformity, fields are represented as methods without parameters (with the exception of the standard self-parameter). The “body” of a field is either a value or yet undefined. Note that the methods are stored in the classes but the fields are kept in the objects, of course. In freshly created objects, the lock is free, and all fields carry the undefined reference ⊥c, where class name c is the (return) type of the field.

We use f for instance variables or fields and l = ζ(s:T).λ().v, resp. l = ζ(s:T).λ().⊥c for field variable definition. Field access is written as v.l() and field update as v’.l := ζ(s:T).λ().v. By convention, we abbreviate the latter constructs by l = v, l = ⊥c, v.l, and v’.l := v. Note that the construct v.l() is used for field access only, but not for method invocation. We will also use v_⊥ to denote either a value v or a symbol ⊥c for being undefined. Note that the syntax does not allow to set a field back to undefined. Direct access (read or write) to fields across object boundaries is forbidden by convention, and we do not allow method update. Instantiation of a new object from class c is denoted by new c.

Expressions especially include syntax to deal with promises and futures. The expression promise T creates a new promise, i.e., a reference or name for a result yet to come. At the point of creation, only the name exists, but no code has been determined and attached to the reference to calculate the result. Binding code to the promise is done by bind o.l(⟨v⟩) : T ← n, stipulating that the eventual result is calculated using the method l of object o with actual parameters ⟨v⟩. Executing the binding operation is also known as fulfilling the promise. Some languages do not allow to independently create a name for the eventual result, i.e., creation and binding are done inseparately by one single command. In that situation, one does not speak of promises, but (just) of futures, even if in the literature, sometimes no distinction is drawn between futures and promises. In a certain way, futures and promises can be seen as two different roles of a reference n: the promise-role means, a client can write to the name using the bind-operation, and the future-role represents the possibility to read back an eventual result using the

3 There will be no result value in case of non-terminating methods.
4 t1; t2 (sequential composition) abbreviates let x:T = t1 in t2, where x does not occur free in t2.
The activity to relinquish and re-grab the lock of the object's lock. Consequently, if the computation of the claimed result needs execution in the object, the threads would be forced to wait until the result is available. This behavior is sketched in Fig. 1. Note the order in which the get-operation is not part of the user-syntax. Both expressions are used to read back the value from a future and the difference in behavior is that get unconditionally attempts to get the value, i.e., blocks until the value has arrived, whereas claim gives up the lock temporarily, if the value has not yet arrived, as explained. This behavior is sketched in Fig. 1.

Before continuing with the type system, let us explain how and why we exclude a specific potential deadlock situation in the semantics of the claim statement (though the language does not generally exclude the presence of deadlocks, i.e., it is possible to write a deadlocking program in the language). Remember that if a thread is about to execute a claim statement in an object, it already owns the object's lock. If the claimed result is not yet available, then the claiming thread blocks. During blocking, if we would not release the lock previously, no other thread could execute in the object, since it would require the object's lock. Consequently, if the computation of the claimed result needs execution in the object, the threads would deadlock. Such deadlocks could not be easily excluded syntactically, since release and grab are only auxiliary syntax, i.e., they cannot be used to write programs, and we do not support checking if a thread already finished its computation. Thus we release the lock before blocking, i.e., waiting for the claimed result, and re-grab the lock after the thread got the result.

2.2. Type system

The language is typed and the available types are given in the following grammar:

\[
T ::= B \mid \text{Unit} \mid [T]^+ \mid [T]^\times \mid [l:U, \ldots, l:U] \mid \langle[l:U, \ldots, l:U]\rangle \mid n
\]

\[
U ::= T \times \cdots \times T \to T
\]
Besides base types $B$ (left unspecified; typical examples are booleans, integers, etc.), Unit is the type of the unit value (). Types $[T]^+$ and $[T]^+-$ represent the reference to a future which will return a value of type $T$, in case it eventually terminates. $[T]^+$ indicates that the promise has not yet been fulfilled, i.e., it represents the write-permission to a promise, which implies read-permission at the same time. $[T]^+$ represents read-only permission to a future. The read/write capability is more specific than read-only, which is expressed by the (rather trivial) subtyping relation generated by $[T]^+ - [T]^+$, accompanied by the usual subsumption rule. Furthermore, $[\_]^+$ acts monotonely, and $[\_]^-$ invariantly wrt. subtyping. When not interested in the access permission, we just write $[T]$.

The name of a class serves as the type for its instances. We need as auxiliary type constructions (i.e., not as part of the user syntax, but to formulate the type system) the type or interface of unnamed objects, written $[l_1; U_1, \ldots, l_k; U_k]$ and the interface type for classes, written $[l_1; C_1, \ldots, l_k; U_k]$. We allow ourselves to write $\bar{T}$ for $T_1 \times \cdots \times T_k$ etc. where we assume the number of arguments match in the rules, and write $\text{Unit} \rightarrow T$ for $T_1 \times \cdots \times T_k \rightarrow T$ when $k = 0$.

We are interested in the behavior of well-typed programs, only, and the section presents the type system to characterize those. As the operational rules later, the derivation rules for typing are grouped into two sets: one for typing on the level of components, i.e., global configurations, and secondly one for their syntactic sub-constituents (cf. also the two different levels in the abstract syntax of Table 1).

Table 2 defines the typing on the level of global configurations, i.e., for “sets” of objects, classes, and named threads. On this level, the typing judgments are of the form

$$\Delta \vdash C : \Theta,$$

where $\Delta$ and $\Theta$ are name contexts, i.e., finite mappings from names (of classes, objects, and threads) to types. In the judgment, $\Delta$ plays the role of the typing assumptions about the environment, and $\Theta$ of the commitments of the component, i.e., the names offered to the environment. Sometimes, the words required and provided interface are used to describe their dual roles. $\Delta$ must contain at least all external names referenced by $C$ and dually $\Theta$ mentions the names offered by $C$. Both contexts constitute the static interface information. A pair $\Delta$ and $\Theta$ of assumption and commitment context with disjoint domains is called well-formed.

The empty configuration $\text{0}$ is well-typed in any context and exports no names (cf. rule T-EMPTY). Two configurations in parallel can refer mutally to each other’s commitments and together offer the (disjoint) union of their names (cf. rule T-PAR).

It will be an invariant of the operational semantics that the identities of parallel entities are disjoint wrt. the mentioned names. Therefore, $\Theta_1$ and $\Theta_2$ in the rule for parallel composition are merged disjointly, indicated by writing $\Theta_1 \oplus \Theta_2$ (analogously for the assumption contexts). Also the treatment of the assumption context requires care wrt. the write permissions. In general, $C_1$ and $C_2$ can rely on the same assumptions that also $C_1 \parallel C_2$ in the conclusion uses, as it represents the environment common to $C_1$ and $C_2$. This, however, does not apply to the write-permissions: if $C_1 \parallel C_2$ do have write permission on a promise $n$, which resides in the environment of $C_1 \parallel C_2$, this is represented as $n:[T]^+-$ in the assumptions of that parallel composition. Due to the linear nature of the write permission, however, the binding $n:[T]^+-$ can occur only in the assumptions of either $C_1$ or of $C_2$ in the two premises of T-PAR. In other words, the assumption context of $C_1 \parallel C_2$ must be split as far as the write permissions to promises are concerned. To capture this intuition, we define:

**Definition 2.1.** Let the symmetric operation $\oplus$ on well-formed name contexts be defined as follows:

\[
\begin{align*}
\text{0} \oplus \Delta &= \Delta \\
{n:[T]^+}, \Delta_1 \oplus n:[T]^+ - , \Delta_2 &= n:[T]^+ - , (\Delta_1 \oplus \Delta_2) \\
n,T, \Delta_1 \oplus n,T, \Delta_2 &= n,T, (\Delta_1 \oplus \Delta_2) \\
\Delta_1 \oplus \Delta_2 &= \text{undefined} & T \neq [T]^+ - \text{ for some } T' \\
\end{align*}
\]

Table 2

Typing (component level).

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<tr>
<th>Rule</th>
<th>Type</th>
<th>Environment Relation</th>
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<tbody>
<tr>
<td>T-EMPTY</td>
<td>$\Delta \vdash 0 : ()$</td>
<td></td>
</tr>
<tr>
<td>T-NCLASS</td>
<td>$\Delta \vdash c(c, O) : (c, T)$</td>
<td>$\Delta \vdash c(O) : T$</td>
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<tr>
<td>T-THREAD</td>
<td>$\Delta \vdash n(t) : (n, [T]^+)$</td>
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<td>$\Delta \vdash n(t) : (n, [T]^+)$</td>
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<td>$\Delta \vdash n(t) : (n, [T]^+)$</td>
</tr>
<tr>
<td>T-THREAD'</td>
<td>$\Delta \vdash n(t) : (n, [T]^+)$</td>
<td>$\Delta \vdash n(t) : (n, [T]^+)$</td>
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<tr>
<td>T-THREAD'</td>
<td>$\Delta \vdash n(t) : (n, [T]^+)$</td>
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<tr>
<td>T-THREAD'</td>
<td>$\Delta \vdash n(t) : (n, [T]^+)$</td>
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<tr>
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<td>$\Delta \vdash n(t) : (n, [T]^+)$</td>
<td>$\Delta \vdash n(t) : (n, [T]^+)$</td>
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<td>$\Delta \vdash n(t) : (n, [T]^+)$</td>
<td>$\Delta \vdash n(t) : (n, [T]^+)$</td>
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<tr>
<td>T-THREAD'</td>
<td>$\Delta \vdash n(t) : (n, [T]^+)$</td>
<td>$\Delta \vdash n(t) : (n, [T]^+)$</td>
</tr>
</tbody>
</table>
checked in an assumption context where the name of the class is available. Note also that the premise of assumption context at that level is of the form context, but additionally a stack-organized, variable context $T-NObj$ of allocated and thus checked in a "parallel" context (cf. again the assumption-commitment rule variable context).

For class types, the rule of instantiation of a class--------------------------------------

value do not change the capabilities (cf. rules $T-Stop$ and $T-Unit$). Note that stop has any type (cf. rule $T-Stop$) reflecting
mean, that a method cannot have a side-effect by exporting read-permission, only. As we may refer to, it is checked in the premise by a method call or a return. capability is available at most once in the program and is not duplicated when passed around. A promise is no longer available

The definition is applied, of course, also to name contexts $\Delta$, used for the commitments. The relations $\leq$ are obviously reflexive, transitive, and antisymmetric.

The definition is applied, of course, also to name contexts $\Delta$, used for the commitments. The relations $\leq$ are obviously reflexive, transitive, and antisymmetric.
the fact that control never reaches the point after stop. Further constructs without side effects are the three expressions to manipulate the monitor locks (suspension, lock grabbing, and lock release), object instantiation (T-NewC), and field update. Wrt. field update in rule T-Fupdate, the reason why the update has no effect on the contexts is that we do not allow fields to carry a type of the form $\langle T \rangle$. This effectively prevents the passing around of write-permissions via fields. The rule T-Let for let-bindings introduces a local scope. The change from $\Delta_1$ to $\Delta_2$ and further from $\Delta_2$ to $\Delta_3$ (and analogously for the $\Gamma$s) reflects the sequential evaluation strategy: first $e$ is evaluated and afterwards $t$. For conditionals, both branches must agree on their pre- and post-$\Delta$-contexts, which typically means, over-approximating the effect by taking the upper bound on both as combined effect. Note that the comparison of the values in T-Cond resp. the check for definedness in T-Cond_⊥ has no side-effect on the contexts. The rule for testing for definedness using undefined (not shown) works analogously.

Table 4 deals with futures, promises, and especially the linear aspect of consuming and transmitting the write-permissions. The claim-command fetches the result value from a future; hence, if the reference $n$ is of type $\langle T \rangle^+$, the value itself carries type $T$ (cf. rule T-Claim). The rule T-Get for get works analogously.

The expression promise $T$ creates a new promise, which can be read or written and is therefore of type $\langle T \rangle^+$. Note, however, that the context $\Delta$ does not change. The reason is that the new name created by promise is hidden by a $\nu$-binder immediately after creation and thus does not immediately extend the $\Delta$-context (see the reduction rule Prom below). The binding of a thread $t$ to a promise $n$ is well-typed if the type of $n$ still allows the promise to be fulfilled, i.e., $n$ is typed by $\langle T \rangle^+$ and not just $\langle T \rangle$. The expression claim dereferences a future, i.e., it fetches a value of type $T$ from the reference of type $\langle T \rangle^+$. Otherwise, the expression has no effect on $\Delta$, as reading can be done arbitrarily many times. As an aside: in rule T-claim, the type of $o$ is not checked, as by convention, the claim-statement must be used in the form claim@$(n,s)$ in the user syntax, where self is the self-parameter of the surrounding methods. Reduction then preserves well-typedness so a re-check here is not needed. Similar remarks apply to the remaining rules. The treatment of get is analogous (cf. rules T-Claim and T-Get). For T-Bind, handing over a promise with read/write permissions as an actual parameter of a method call, the caller loses the right to fulfill the promise. Of course, the caller can only pass the promise to a method which assumes read/write permissions, if itself has the write permission. The loss of the write-permission is specified by setting $\Delta$ and $\Gamma$ to $\Delta \setminus n: \langle T \rangle^+$ resp. to $\Gamma \setminus \nu : T$. The difference-operator $\Delta \setminus n: \langle T \rangle^+$ removes the write-permission for $n$ from the context $\Delta$. In T-Bind\_\setminus, the premise $\Gamma; \Delta, n: \langle T \rangle^+ \vdash \nu : T$ abbreviates the following: assume $\nu = \nu_1, \ldots, \nu_n$ and $T = T_1 \ldots T_n$ and let $\Xi_1$ abbreviate

<table>
<thead>
<tr>
<th>$\Gamma; \Delta \vdash e: \langle T \rangle_1 \ldots \langle T \rangle_n$</th>
<th>$\Gamma; \Delta \vdash m_1: U_1 \ldots m_k: U_k$</th>
<th>$\Gamma; \Delta \vdash f_1: U_1 \ldots f_k: U_k$</th>
<th>$\Gamma; \Delta \vdash f_i: U_i$</th>
</tr>
</thead>
<tbody>
<tr>
<td>$\Gamma; \Delta \vdash \lambda x_1 \ldots x_n : \langle T \rangle_1 \ldots \langle T \rangle_n \cdot t$</td>
<td>$\Gamma; \Delta \vdash x_1 \ldots x_n : \langle T \rangle_1 \ldots \langle T \rangle_n$</td>
<td>$\Gamma; \Delta \vdash t : \langle T \rangle_1 \ldots \langle T \rangle_n$</td>
<td>$\Gamma; \Delta \vdash t : \langle T \rangle_1$</td>
</tr>
<tr>
<td>$\Gamma; \Delta \vdash f_i : \langle T \rangle_1 \ldots \langle T \rangle_n$</td>
<td>$\Gamma; \Delta \vdash f_i : \langle T \rangle_1$</td>
<td>$\Gamma; \Delta \vdash f_i : \langle T \rangle_1$</td>
<td>$\Gamma; \Delta \vdash f_i : \langle T \rangle_1$</td>
</tr>
<tr>
<td>$\Gamma; \Delta \vdash t : \langle T \rangle_1 \ldots \langle T \rangle_n$</td>
<td>$\Gamma; \Delta \vdash t : \langle T \rangle_1$</td>
<td>$\Gamma; \Delta \vdash t : \langle T \rangle_1$</td>
<td>$\Gamma; \Delta \vdash t : \langle T \rangle_1$</td>
</tr>
</tbody>
</table>

<table>
<thead>
<tr>
<th>$\Gamma; \Delta \vdash e : \langle T \rangle_1 \ldots \langle T \rangle_n$</th>
<th>$\Gamma; \Delta \vdash v : \langle T \rangle_1 \ldots \langle T \rangle_n$</th>
<th>$\Gamma; \Delta \vdash e : \langle T \rangle_1 \ldots \langle T \rangle_n$</th>
<th>$\Gamma; \Delta \vdash e : \langle T \rangle_1$</th>
</tr>
</thead>
<tbody>
<tr>
<td>$\Gamma; \Delta \vdash e_1 : \langle T \rangle_1$</td>
<td>$\Gamma; \Delta \vdash e_1 : \langle T \rangle_1$</td>
<td>$\Gamma; \Delta \vdash e_1 : \langle T \rangle_1$</td>
<td>$\Gamma; \Delta \vdash e_1 : \langle T \rangle_1$</td>
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<td>$\Gamma; \Delta \vdash e_1 : \langle T \rangle_1$</td>
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</tbody>
</table>

$\Gamma; \Delta \vdash e : \langle T \rangle_1 \ldots \langle T \rangle_n$ $\Gamma; \Delta \vdash e : \langle T \rangle_1$ $\Gamma; \Delta \vdash e : \langle T \rangle_1$ $\Gamma; \Delta \vdash e : \langle T \rangle_1$
Table 4
Typing (objects and threads).

<table>
<thead>
<tr>
<th>Rule</th>
<th>Description</th>
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</thead>
<tbody>
<tr>
<td>$\Gamma; \Delta \vdash \text{promise } T : [T]^+$</td>
<td>T-PROM</td>
</tr>
<tr>
<td>$\Gamma; \Delta \vdash n : [T]^+$</td>
<td>T-CLAIM</td>
</tr>
<tr>
<td>$\Gamma; \Delta \vdash \alpha \cdot c$</td>
<td>T-GET</td>
</tr>
<tr>
<td>$\Gamma; x : T \vdash \Delta$</td>
<td>T-VAR</td>
</tr>
<tr>
<td>$\Delta = \Delta \setminus n : T$</td>
<td>T-NAME</td>
</tr>
<tr>
<td>$\Gamma(x) = T, ; \Gamma = \Gamma \setminus x : T$</td>
<td>T-BOUND</td>
</tr>
<tr>
<td>$\Delta = \Delta \setminus {n : T}$</td>
<td>T-SUSPEND</td>
</tr>
<tr>
<td>$\Delta = \Delta \setminus {c}$</td>
<td>T-GRAB</td>
</tr>
<tr>
<td>$\Gamma_1 \vdash \text{suspend}(o) : \text{Unit}$</td>
<td>T-RELEASE</td>
</tr>
</tbody>
</table>

$\Gamma_1; \Delta \vdash t : T \Rightarrow \Gamma_2; \Delta_2 \; \text{T-SUB}$

$\Gamma; \Delta \vdash n : [T]^+$.

For completeness, we also list the typing rules for objects and threads. These rules are used to ensure that objects and threads are well-typed.

\[ \Gamma; \Delta \vdash n : [T]^+ \]

\[ \Gamma; \Delta \vdash \alpha \cdot c \]

\[ \Gamma; \Delta \vdash \text{suspend}(o) : \text{Unit} \]

\[ \Gamma; \Delta \vdash \text{grab}(o) : \text{Unit} \]

2.3. Operational semantics

The operational semantics is given in two stages, component internal steps and external ones, where the latter describe the interaction between component and environment. They are defined in reference to assumption and commitment contexts. The static part of the contexts corresponds to the static type system from Section 2.2 on component level and takes care that, e.g., only well-typed values are received from the environment.

2.3.1. Internal steps

The internal semantics describes the operational behavior of a closed system, not interacting with its environment. The corresponding reduction steps are shown in Table 5, distinguishing between confluent steps $\rightsquigarrow$ and other internal transitions $\Rightarrow$, both invisible at the interface. The $\rightsquigarrow$-steps, on the one hand, do not access the instance state of the objects. They are free of imperative side effects and thus confluent. The $\Rightarrow$-steps, in contrast, access the instance state, either by reading or by writing it, and may thus lead to race conditions. In other words, this part of the reduction relation is in general not confluent.

The first seven rules deal with a basic sequential constructs, all as $\rightsquigarrow$-steps. The basic evaluation mechanism is substitution (cf. rule $\text{RED}$). Note that the rule requires that the leading let-bound variable is replaced only by values $v$. The operational behavior of the two forms of conditionals are axiomatized by the four $\text{COND}$-rules. Depending on the result of the comparison in the first pair of rules, resp., the result of checking for definedness in the second pair, either the then- or the else-branch is taken. In $\text{COND}_2$, we assume that $v_1$ does not equal $v_2$, as side condition. Evaluating stop terminates the future for good, i.e., the rest of the thread will never be executed as there is no reduction rule for the future n(stop) (cf. rule $\text{STOP}$). The rule $\text{LOOKUP}$ deals with field look-up, where $F.\ell(o)$ stands for $\ell_{c}(o[s]) = \ell_{c}$, resp., for $\forall \ell(o/s)$, where $[c, F, L] = [c, \ldots, l = \xi(s), c) \lambda(\ldots, L]$, if the field is yet undefined, resp., $[c, F, L] = [c, \ldots, l = \xi(s,c), c) \lambda(\ldots, L]$. In $\text{FUPDATE}$, the meta-mathematical notation $F.\ell := v$ stands for $(\ldots, l = v, \ldots)$, when $F = (\ldots, l = v, \ldots)$. There will be no external variant of the rule for field look-up later in the semantics of open systems, as we do not allow field access across component boundaries. The same restriction holds for field update in rule $\text{FUPDATE}$. A new object as instance of a given class is created by rule $\text{NEWObj}$. Note that initially, the lock is free and there is no activity associated with the object, i.e., the object is initially passive.

The expression promise $T$ creates a fresh promise $n'$. A new thread $n'(\bullet)$ is allocated with an “undefined” body, as so far nothing more than the name is known. The rule $\text{PROM}$ mentions the types $T$ and $T'$. The typing system assures that the type $T$ is of the form $[S]^+$ for some type $S$. A promise is fulfilled by the bind-command (cf. rule $\text{BIND}_1$), in that the new thread $n'$ is
put together with the code to be executed and run in parallel with the rest. In the configuration after the reduction step, the meta-mathematical notation $M.o(\bar{v})$ stands for $(o[s]/\bar{v}[\bar{x}])$, when the method suite $[M]$ equals $\ldots l = \ldots$.

Upon termination, the result is available via the claim- and the get-syntax (cf. the CLAIM-rules and rule GET1), but not before the lock of the object is given back again using release(o) (cf. rule RELEASE). If the thread is not yet terminated, the requesting thread suspends itself, thereby giving up the lock. The behavior of claim is sketched in Fig. 1. Note the types of the involved let-bound variables: the future reference is typed by $[T]$, indicating that the value for x will not directly be available, but must be dereferenced first via claim.

The two operations grab and release take, resp., give back an object’s lock. They are not part of the user syntax, i.e., the programmer cannot directly manipulate the monitor lock. The user can release the lock using the suspend-command or by trying to get back the result from a call using claim.

The above reduction relations are used modulo structural congruence, which captures the algebraic properties of parallel composition and the hiding operator. The basic axioms for $\equiv$ are shown in Table 6 where in the fourth axiom, $n$ does not occur free in $C_1$. The congruence relation is imported into the reduction relations in Table 7. Note that all syntactic entities are always tacitly understood modulo $\alpha$-conversion.

For illustration of the operational semantics, we show the combination of creating a promise and binding a method body to it. The steps in the reduction sequence below are justified by PROM, LET, and BIND, in that order. In the sequence, we did not write the definition of the object plus the class, needed to do the last reduction step. I.e., the reduction sequence below runs in parallel with $c[F,M] \equiv c[F,L]$ where in particular the method suite $M$, stored in the class $c$ of the object $o$, contains the definition of the method body. That definition is needed for binding operation in the last reduction step. In the corresponding rule BIND1, this is written as $M.o(\bar{v})$. In the final configuration, $t'$ contains the result of looking up the method body and is of the form $\text{get}(o); M.o(\bar{v})$. This, the overall behavior of the fulfilled promise $n_2$, i.e., after the binding step, is: first acquire the lock of the object, afterwards executed the method body with the formal parameters including the
First we characterize as erroneous situations where a promise is about to be written a second indeed fulfilled only once. An important part of it is a standard property, namely preservation of well-typedness under internal generalizes the setting with futures and asynchronous method calls, only. asynchronous method calls can therefore be seen as syntactic sugar. The introduction of promises as a separate datatype and binding as corresponding, separate operation on promises therefore generalizes the setting with futures and asynchronous method calls, only.

In the following, we show that a well-typed component does not contain a manifest write-error. The necessary ancillary lemmas will in general proceed by induction on the typing derivations for judgments of the form \( \Delta \vdash C : \Theta \). From a proof-theoretical (and algorithmic) point of view, the type system as formalized in Tables 2, 3, and 4 has an unwelcome property: it is too "non-deterministic" in that it allows the non-structural subsumption rules T-Sub on the level of threads \( t \) and on the level of components \( C \) at any point in the derivation. This liberality is unwelcome for proofs by induction on the typing derivation as one loses knowledge about the structure of the premises of an applied rule in the derivation. We write \( \Delta \vdash_m C : \Theta \) for derivations where subsumption at the level of components (by rule T-Sub from Table 2) is not used, and subsumption from Table 4 is only used "when needed", i.e., for adaptation. Taking for instance T-Bind and concentrating on the premises relevant for the illustration: Given as the interface type of the class \( \Gamma; \Delta, n: [T]^+ \vdash_m C : \{[\ldots, \ell T \rightarrow T, \ldots]\} \) and furthermore \( \Gamma; \Delta, n: [T]^+ \vdash_m \nu ; \hat{S} \), the minimal types \( \hat{S} \) of the \( \nu \) may not directly match the expected argument type \( T \) of the method labeled \( \ell \) (as is required in the premise of the rule T-Bind). Restricting now the use of subsumption to "adapt" the \( \hat{S} \) to \( T \) gives the type system for minimal types (denoted by using \( \vDash_m \) instead of \(+\)). This could be explicitly done by removing the freely applicable T-Sub and distributing its effect into the premises of structural rules, where such adaptation is needed.

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In the discussed rule T-Bind, by stipulating

\[
\Gamma; \Delta, n: [T]^+ \vdash_m \nu ; \hat{S} \leq \bar{T} \tag{\text{T-Bind}_m}
\]

where \( \hat{S} \leq \bar{T} \) is interpreted pointwise \( S_i \leq T_i \), for all \( i \). As the formulation of that type system is rather standard and straightforward, we omit its definition.

**Lemma 2.3** (Minimal typing).
1. If \( \Delta \vdash_m C : \Theta \) and \( \Delta' \vdash C : \Theta' \), then \( \Delta \leq \Delta' \) and \( \Theta \leq \Theta' \).
2. If \( \Delta \vdash_m C : \Theta \) then \( \Delta \vdash C : \Theta \).
3. If \( \Delta' \vdash C : \Theta' \), then \( \Delta \vdash_m C : \Theta \) with \( \Delta \leq \Delta' \) and \( \Theta \leq \Theta' \), for some \( \Delta \) and \( \Theta \).

**Proof.** Straightforward. \( \square \)

First we show that a well-typed component does not contain a manifest write-error.

**Lemma 2.4.** If \( \Delta \vdash_m C : \Theta \), then \( \vdash C : ok \).
Proof. By induction on the typing derivations for judgments on the level of components, i.e., for judgments of the form $\Delta \vdash C : \Theta$; the subordinate typing rules from Tables 3 and 4 on the level of threads and expressions do not play a role for the proof. The empty component in the base case of T-EMPTY is clearly write-error free. The cases for the T-Nu-rules by straightforward induction. The cases for T-NCLASS, T-NOBJ, and T-NFUTURE are trivially satisfied, as they mention a single, basic component, only.

Case: T-Par
We are given $\Delta_1, \Theta_2 \vdash C : \Theta_1$ and $\Delta_2, \Theta_1 \vdash C : \Theta_2$ with $\Delta = \Delta_1 \uplus \Delta_2$. By induction, both $C_1$ and $C_2$ are write-error free. The non-trivial case (which we will lead to a contradiction) is when one of the components attempts to write to a promise and the partner already has fulfilled it. So, without loss of generality assume that $C_1 = \nu(\Theta_1), (C'_1 \parallel n_1 \mid \text{let } x : T = \text{bind } x : T \leftarrow n_2 \in t^+) \parallel C_2 = \nu(\Theta_2). (C'_2 \parallel n_2(t_2)))$. Assume that $n_2$ occurs in neither $\Theta'_1$ nor $\Theta'_2$, otherwise no write error is present (since in that case, the name $n_2$ mentioned on both sides of the parallel refer to different entities). For $C_1$ to be well-typed, we have $\Delta_1, \Theta_2 \vdash n_2 : [T_2]^+ \rightarrow$ for some type $T_2$. For $C_2$ to be well-typed, we have $\Theta_2 \vdash n : [T_2]^+$ for some type $T_2$. Thus, $\Delta \vdash C_1 \parallel C_2 : \Theta_1, \Theta_2$ cannot be derived, which contradicts the assumption. \hfill $\Box$

Lemma 2.5 (Subject reduction: $\Downarrow$). If $\Delta \vdash_m C_1 : \Theta$ and $C_1 \equiv C_2$, then $\Delta \vdash_m C_2 : \Theta$.

Proof. We show preservation of typing by the axioms of Table 6. Proceed by induction on the derivation of $\Delta \vdash_m C_1 : \Theta$.

Case: $C \equiv \parallel \Theta$ (idempotence)
We are given $\Delta \vdash C \equiv \parallel \Theta$. Inverting T-Par and by T-EMPTY we get as sub-goals $\Delta, \Theta \vdash \parallel \Theta \equiv ()$ and $\Delta \vdash_m C : \Theta$, which concludes the case.

Case: $C_1 \equiv C_2 \equiv C_1$ (commutativity)
Immediate using T-Par and T-EMPTY.

Case: $C_1 \equiv (C_2 \parallel C_3) \equiv (C_1 \parallel C_2) \equiv C_2$ and vice versa (associativity)
By straightforward induction.

Case: $C_1 \equiv \nu(n:T), C_2 \equiv \nu(n:T). (C_1 \parallel C_2)$
where $n$ does not occur free in $C_1$. We are given $\Delta \vdash C_1 \parallel \nu(n:T), C_2 : \Theta_1, \Theta_2$, where $n$ occurs in neither $\Theta_1$ or $\Theta_2$. Inverting T-Par and T-Nu, we obtain as two subgoals $\Delta, \Theta_2 \vdash C_1 : \Theta_1$ and $\Delta, \Theta_1 \vdash C_2 : \Theta_1, \Theta_2, n:T$, and the result follows by T-Par and the T-Nu-rule.

Case: $\nu(n_1:T_1), \nu(n_2:T_2). C \equiv \nu(n_2), \nu(n_1:T_1), C$
Analogously. \hfill $\Box$

The next lemma is another step towards subject reduction. Note that minimal types are not preserved by reduction. Especially executing a bind-operation with rule BIND, changes the type of the corresponding name from $[T]^+ \rightarrow$ to $[T]^+$.

Lemma 2.6 (Subject reduction: $\Downarrow$ and $\Downarrow$). Assume $\Delta \vdash_m C : \Theta$.

1. If $C \Downarrow \hat{C}$, then $\Delta \vdash \hat{C} : \Theta$.

2. If $C \Downarrow \Downarrow$, then $\Delta \vdash \Downarrow : \Theta$.

Proof. The reduction rules of Table 5 are all of the form $C_1 \parallel n(t_1) \Downarrow \Downarrow C_2 \parallel n(t_2)$, where often $C_1 = C_2$ or $C_1$ and $C_2$ missing. In the latter case, it suffices to show that $\Delta, n:[T]^+ \vdash_m t_1 : T \implies \Delta, n:[T]^+ \vdash t_2 : T$.

Case: Red: $n\parallel \text{let } x : T = v \in t \Downarrow \Downarrow n\parallel v[x/v]$ 
By preservation of typing under substitution.

The 5 rules for let and for conditionals are straightforward. The case for stop follows from the fact that stop has every type (cf. rule T-Stop).

Case: Prom: $n\parallel \text{let } x : T' = \text{promise } T \in t \Downarrow \Downarrow n\parallel \text{let } x : T' = n' \in t) \parallel n'(\bullet)$
The type system (for minimal types) assures that $T' = [T]^+$, i.e., for the left-hand side of the reduction step, we obtain as one subgoal (inverting T-NTHREAD', T-LET, and T-PROM) $x:[T]^+ \vdash \Delta, n:[S]^+ \vdash t : S$. The result follows from T-Nu, T-Par, T-Let, and T-NTHREAD' (and weakening):
Case: \( \text{BIND}_1 \) \( n_1(t) \parallel n_2(\vec{v}) = n_1(\text{let} x:T = \text{bind} o.l(\vec{v}):T \leftarrow n_2 \in t_1) \parallel n_2(\text{let} x:T = \text{grab}(o); M.i(o)(\vec{v}) \text{in release}(o); x) : n_2[T]^{++} \)

The type system assures (cf. T-BIND) that \( T = [T_2]^{++} \). By assumption, we are given \( \Delta \vdash n_1(t) : \Theta \), which implies \( \Theta = n_1[T_1]^{++} \) for some type \( T_1 \). Inverting rules T-PAR, T-NTHREAD, T-LET, and T-BIND gives for the named thread \( n_1 \):

\[
\Delta, n_2[T_2]^{++}, n_1[T_1]^{++} \vdash \vec{v} : \vec{v}' = \Delta'' = \Delta' : (\vec{v}, \vec{v}') \quad \text{T-BIND}
\]

\[
\Delta, n_2[T_2]^{++}, n_1[T_1]^{++} \vdash \text{let} x:T = \text{bind} o.l(\vec{v}):T \leftarrow n_2 \in t_1 : T_1 = : \Delta, n_2[T_2]^{++}, n_1[T_1]^{++}
\]

Rule T-BIND (and T-NTHREAD) implies that the assumption context \( \Delta \) contains especially the binding \( n_2[T]^{++} \), i.e., the assumption \( \Delta \) in the last conclusion is of the form \( \Delta', n_2[T_2]^{++} \).

Now to the post-configuration after the \( \downarrow \) step. With T-PAR we obtain the following two sub-goals:

\[
\Delta, n_2[T_2]^{++}, n_1[T_1]^{++} \vdash n_2(\text{let} x:T = n_2 \in t_1) : n_1[T_1]^{++} \quad \Delta, n_1[T_1]^{++} \vdash n_2(\text{let} x:T = \text{grab}(o); M.i(o)(\vec{v}) \text{in release}(o); x) : n_2[T_2]^{++}
\]

The left one can be derived using T-NTHREAD, T-LET, and T-NAME, where there second premise of T-LET is discharged by the corresponding assumption from above and weakening

\[
\Delta, n_2[T_2]^{++}, n_1[T_1]^{++} \vdash n_2 : [T_2]^{++} \quad \Delta, n_2[T_2]^{++}, n_1[T_1]^{++} \vdash \text{let} x:T = n_2 \in t_1 : T_1
\]

The second premise can be derived as follows:

\[
\Delta, n_1[T_1]^{++}, n_2[T_2]^{++} \vdash M.i(o)(\vec{v}) : T_2 \quad \Delta, n_1[T_1]^{++}, n_2[T_2]^{++} \vdash \text{let} y:T = M.i(o)(\vec{v}) \text{in release}(o); y : T_2
\]

The premise \( \Delta, n_1[T_1]^{++}, n_2[T_2]^{++} \vdash M.i(o)(\vec{v}) : T_2 \) follows by preservation of typing by substitution. Note the use of subsumption in the last step.

Lemma 2.7 (Subject reduction: \( \equiv \)). If \( \Delta \vdash C_1 : \Theta \) and \( C_1 \equiv C_2 \), then \( \Delta \vdash C_2 : \Theta \).

Proof. Assume \( \Delta \vdash C_1 : \Theta \) and \( C_1 \equiv C_2 \). By Lemma 2.3(3), \( \Delta' \vdash_m C_1 : \Theta' \) s.t. \( \Delta \leq \Delta' \) and \( \Theta' \leq \Theta \). By Lemma 2.5, \( \Delta' \vdash_m C_2 : \Theta' \), and hence by Lemma 2.3(2), also \( \Delta' \vdash C_2 : \Theta' \), and the result follows by subsumption (rule T-SUB).
Labels.  

<table>
<thead>
<tr>
<th>( \gamma )</th>
<th>( \text{basic labels} )</th>
</tr>
</thead>
<tbody>
<tr>
<td>( \gamma ::= n(\text{call } o.l(\theta)) \mid n(\text{get}(v)) \mid v(n:T).\gamma )</td>
<td>receive and send labels</td>
</tr>
</tbody>
</table>

**Lemma 2.8** (Subject reduction: \( \Rightarrow \) and \( \Rightarrow \)). Assume \( \Delta \vdash C : \Theta \).
1. If \( C \Rightarrow \hat{C} \), then \( \Delta \vdash \hat{C} : \Theta \).
2. If \( C \Rightarrow \hat{C} \), then \( \Delta \vdash \hat{C} : \Theta \).

**Proof.** As consequence of the corresponding property for minimal typing from Lemma 2.6, Lemma 2.3, and subsumption. \( \square \)

**Lemma 2.9** (Subject reduction). If \( \Delta \vdash C : \Theta \) and \( C \Rightarrow \hat{C} \), then \( \Delta \vdash \hat{C} : \Theta \).

**Proof.** A consequence of Lemma 2.7 and 2.8. \( \square \)

A direct consequence is that all reachable configurations are write-error free:

**Corollary 2.10.** If \( \Delta \vdash C : \Theta \) and \( C \Rightarrow \hat{C} \), then \( \Delta \vdash \hat{C} : \text{ok} \).

**Proof.** A consequence of Lemma 2.4 and subject reduction from Lemma 2.9. \( \square \)

2.3.2. External semantics

In this section we introduce the external semantics that defines the interaction between component and environment. We start by formalizing typing judgments and transitions between typing judgments, being the basic form of the external steps. We continue with static typing assumptions for well-formed and well-typed labels. Context updates, given next, express the dynamic change of typing judgments for incoming and outgoing communications. Making use of the above formalisms, we give the steps of the external semantics.

The external semantics formalizes the interaction of an open component with its environment. The semantics is given as labeled transitions between typing judgments on the level of components (cf. Table 2), i.e., judgments of the form

\[ \Delta \vdash C : \Theta, \]

where, as before, \( \Delta \) represents the assumptions about the environment of the component \( C \) and \( \Theta \) the commitments. The assumptions require the existence of named entities in the environment (plus giving static typing information), and dually, the commitment promises the existence of such entities in \( C \). It is an invariant of the semantics, that the assumption and commitment contexts are disjoint concerning their name bindings. In addition, the interface keeps information about whether the value of a future \( n \) is already known at the interface (this is a bit of information not needed in the static type system of Table 2). If it is, we write \( n:T = v \) as binding of the context. We write furthermore \( \Delta \vdash n = v \), if \( \Delta \) contains the corresponding value information (and if not interested in the type) and write \( \Delta \vdash n = \bot \), if that is not the case. This extension makes the value of a future (once successfully claimed) available at the interface. With these judgments, the external transitions are of the form:

\[ \Delta \vdash C : \Theta \quad \xrightarrow{\Delta} \quad \Delta \vdash \hat{C} : \Theta. \]

**Notation 2.11.** We abbreviate the tuple of name contexts \( \Delta, \Theta \) as \( \Xi \). Furthermore we understand \( \hat{\Delta}, \hat{\Theta} \) as \( \hat{\Xi} \), etc.

The labels of the external transitions represent single steps of the interface interactions (cf. Table 8).

A component exchanges information with the environment via \textit{call} and \textit{get} labels (by convention, referred to as \( \gamma_c \) and \( \gamma_g \), for short). Interaction is either incoming or outgoing, indicated by \( ? \), resp., \( ! \). In the labels, \( n \) is the identifier of the thread (i.e., also future/promise) carrying out the call resp. of being queried via claim or get. Besides that, object and future names (but no class names) may appear as arguments in the communication. Scope extrusion of names across the interface is indicated by the \textit{v}-binder. Given a basic label \( \gamma = v(\Xi).\gamma' \) where \( \Xi \) is a name context such that \( v(\Xi) \) abbreviates a sequence of single \( n:T \) bindings (whose names are assumed all disjoint, as usual) and where \( \gamma' \) does not contain any binders, we call \( \gamma' \) the \textit{core} of the label and refer to it by \( [\gamma] \). We define the core analogously for receive and send labels. The free names \( fn(a) \) and the bound names \( bn(a) \) of a label \( a \) are defined as usual, whereas \textit{names}(a) refer to all names of \( a \). In addition, we distinguish between names occurring as arguments of a label, in \textit{passive} position, and the name occurring as carrier of the activity, in \textit{active} position. Name \( n \), for illustration, occurs actively and free in \( n(\text{call } o.l(\bar{v})) \) and in \( n(\text{get}(v)) \). We write \( fn_p(a) \) for the free names occurring in active position, \( fn_p(a) \) for the free names in passive position, etc. All notations are used analogously for...
and if a future value is claimed successfully for the first time. For outgoing communication, the situation is dual. Conversely especially extended by the new names, whose scope extrudes. For the binding part Typechecking labels.

For incoming calls, the commitment context \( \Gamma \) is simpler. Hence, we concentrate here on basic labels \( y \). Note that for incoming labels, \( \Xi \) contains only bindings to environment objects (besides future names), as the environment cannot create component objects; dually for outgoing communication.

The steps of the operational semantics for open systems check the well-formedness and well-typedness of the call is checked using only the commitment context \( \Gamma \). Besides information about new names, the context information is potentially updated at the same time. The information \( \delta \) is interpreted in such a way that checking for write-permission \( \delta \) is no longer available for type checking the rest of the sequence.

In a similar spirit: requiring that \( \delta \) is defined dually. In particular, in the first premise, \( \hat{\delta} \) is replaced by \( \hat{\delta} \). Well-typedness of an incoming core label \( a \) with expected type \( \vec{T} \), resp., \( T \), and relative to the name context \( \Xi \) is asserted by

\[
\hat{\Xi} \vdash a : \vec{T} \rightarrow \_ \quad \text{resp.,} \quad \hat{\Xi} \vdash a : \_ \rightarrow T,
\]

as given by Table 9. Finally, let \( \hat{\Xi}_0 \) abbreviate \( \hat{\Xi} \). Then \( \hat{\Xi} \vdash \vec{v} : \vec{T} \) means: \( \hat{\Xi}_i \vdash v_i : T_i \) and \( \hat{\Xi}_{i+1} = \hat{\Xi}_i \setminus T_i \), for all \( 0 \leq i \leq n - 1 \).

Note that the receiver \( o \) of the call is checked using only the commitment context \( \delta \), to assure that \( o \) is a component object. Note further that to check the interface type of the class \( c \), the full \( \Xi \) is consulted, since the argument types \( \vec{T} \) or the result type \( T \) may refer to both component and environment classes.

The premise \( \hat{\Xi} \vdash \vec{v} : \vec{T} \) in LT-CALL is interpreted in such a way that checking for write-permission consumes that permission (analogous to the corresponding premise of T-BIND in Table 4). This is formalized in the definition of \( \hat{\Xi} \vdash \vec{v} : \vec{T} \) for well-typedness of a sequence of values, given at the end of Definition 2.12, which iterates through the sequence, potentially removing write-permission for a \( v_i \) s.t. the permission is no longer available for type checking the rest of the sequence.

In a similar spirit: requiring that \( \Xi \) is of the form \( \Xi_1, n;\{T\}^+ \). \( \hat{\Xi}_2 \) assures that it is not possible to transmit \( n \) with write-permissions if \( n \) is the active thread of the label.

Besides checking whether the assumptions are met before a transition, the contexts are updated by a transition step, especially extended by the new names, whose scope extrudes. For the binding part \( \Xi' \) of a label \( o(\Xi'), y \), the scope of the references to existing objects and name labels \( \Lambda' \) extrudes across the border. In the step, \( \Lambda' \) extends the assumption context \( \Delta \) and \( \Theta' \) the commitment context \( \Theta \). Besides information about new names, the context information is potentially updated wrt. the availability of a future value. This is done when a get-label is exchanged at the interface for the first time, i.e., when a future value is claimed successfully for the first time. For outgoing communication, the situation is dual.

Before we come to the corresponding Definition 2.13 below, we make clear (again) the interpretation of judgments \( \Delta \vdash C : \Theta \). Interesting is in particular the information \( n;\{T\}^+ \), stipulating that name \( n \) is available with write-permission (and result type \( T \)). In case of \( \Delta \vdash n : \{T\}^+ \), the name \( n \) is assumed to be available in the environment as writable, and conversely \( \Theta \vdash n : \{T\}^+ \) asserts write permission for the component. Since read permissions, captured by types \( \{T\}^+ \), are not treated linearly – one is allowed to read from a future reference as many times as wished – the treatment of bindings \( \{n;T\}^+ \) is simpler. Hence, we concentrate here on \( n;\{T\}^+ \) and the write permissions.

As the domains of \( \Delta \) and \( \Theta \) are disjoint, bindings \( n; T \) cannot be available in the assumption context \( \Delta \) and the commitments \( \Theta \) at the same time. The information \( T' = \{T\}^+ \) indicates which side, component or environment, has the write permission. If, e.g., \( \Delta \vdash n : \{T\}^+ \), then the component is not allowed to execute a claim-operation on \( n \). In the mentioned situation, the component can execute a claim-operation on \( n \). The same applies if \( \Delta \vdash n : \{T\}^+ \). In other words, a name \( n \) can be accessed
by reading by both the environment and the component once known at the interface, independent from whether it is part of Δ or of Θ. A difference between bindings of the form n: [T]⁺ and n: [T]⁻ (and likewise n: [T]⁺ = v) is, that communication can change Δ ⊢ n : [T]⁺ to Θ ⊢ n : [T]⁻ and vice versa. For names n of type [T]⁺, this change of side is impossible. The latter kind of information, for instance Θ ⊢ n : [T]⁺, implies that the code has been bound to n and it is placed in the component. Once fixed there, the reference to n may, of course, be passed around, but the thread named n itself cannot change to the environment since the language does not support mobile code.

Now, how do interface interactions update the contexts? We distinguish two ways, the name n can be transmitted in a label: passively, when transported as the argument of a call or a get-interaction, and actively, when mentioned as the carrier of the activity, as the n in n(call o.l(v)) and n(get(v)). As usual, such references (actively or passively) can be transmitted as fresh names, i.e., under a ν-binder, or alternatively as an already known name. When transmitted passively and typed with [T]⁻ for some type T, the write-permission to n is handed over to the receiving side and at the same time, that permission is removed from the sender side.

Now, what about transmitting n actively? An incoming call n(call o.l(v))? e.g., reveals at the interface that the promise indeed has been fulfilled. As, in that situation of an incoming call, the thread, executing the call, is located at the component, the commitment context is updated to satisfy Θ ⊢ n : [T]⁻ = ⊥ (for an appropriate type T) after the communication. Indeed, before the step it is checked, that the environment actually has write permission for n, i.e., that Δ ⊢ n : [T]⁺, or that the name n is new. See the incoming call in Fig. 2a, where the n is fresh, resp. in Fig. 2c, where the n has been transmitted passively and with write-permissions to the environment before the call (in the dotted arrow).

Whereas call-labels make public, at which side the thread in question resides, get-labels, on the other hand, reveal that the computation has terminated and fix the result value (if the information about the result value had not been public interface information before). There are two situations, where a, say, outgoing get-communication is possible. In both cases, the named thread, representing the future, resides in the component and after the get-communication, the value is determined, i.e., Θ ⊢ n : [T]⁺ = v (if not already before the step). One scenario is that Δ ⊢ n : [T]⁺ = ⊥ before the step still. If, in that situation, the get is executed by the environment, it is required that the component must have had write permission before, i.e., Θ ⊢ n : [T]⁺ (cf. Fig. 2b). The only way, the value for n is available for the environment now is that the promise had been fulfilled and the corresponding thread already has terminated, and this could have been done by the component, only. In that situation, the contexts are updated from Θ ⊢ n : [T]⁺ to Θ ⊢ n : [T]⁻ = v by the get-interaction. Alternatively, the thread may be known to be part of the component with the promise already fulfilled (Θ ⊢ n : [T]⁻ = ⊥, as shown in Fig. 2a and c). Finally, the value for n might already been known at the interface, i.e., already before the step, Θ ⊢ n : [T]⁻ = v holds. In that situation, v has been added as interface information previously by a prior get-interaction, and the situation corresponds to the very last get in Fig. 2b and c.

**Definition 2.13** (Context update). Let Ξ be a name context and a = ν(Ξ). [a] an incoming label. Let ˘Ξ = Ξ + a be defined asfollows.

We define the (intermediate) contexts Θ'' = Θ and Δ'' = Δ, ˘Ξ. Let furthermore Σ'' be the set of bindings defined as follows. In case of a call label, i.e., [a] = n(call o.l(v))? let the vector of types T be defined by Ξ ⊢ o.l : T → T according to Equation (5) of Definition 2.12. Then Σ'' consists of bindings of the form ν(Ξ) for values ν from v such that T_i = [T_i]⁺⁻. In case of a get label, i.e., [a] = n(get(v))? the context Σ'' is ν(Ξ) if Δ'' ⊢ n : [([T]⁻)⁺⁻], and empty otherwise.

With Σ'' given this way, the definitions of the post-contexts Δ and Θ distinguish between calls and get-interaction: If a is a call label and n ∈ names_o(a), we define

\[ \triangledown = (\Delta'' \setminus \Sigma'') \setminus n: [T]⁻ \quad \text{and} \quad \Theta'' = \Sigma'', n: [T]⁺. \]  

(8)

If a is a get label a = ν(Ξ). n(get(v))? and n ∈ names_o(a), Δ and Θ are given by:

\[ \triangledown = (\Delta'' \setminus \Sigma''), n: [T]⁺ = v \quad \text{and} \quad \Theta'' = \Sigma'', \]  

(9)
For **outgoing** communication, the definition is applied dually.

The definition proceeds in two stages. In a first step, the assumption context $\Delta$ is extended with the bindings $\Sigma'$ carried with the incoming label $a$. The second step deals with the write permissions, i.e., it transfers the write permission transmitted from the sender to the receiver. The binding context $\Sigma''$ deals with the permissions carried by thread names transmitted passively, i.e., as arguments of the communication. It remains to take care also of the information carried by the active thread. There, we distinguish calls and get-labels. An incoming call (Equation (8)) with $n$ as active thread is the sign that the thread is now located at the component side and that the write permission has been consumed by the environment. Hence, in Equation (8), the environment loses the write-permission and the component is extended by the binding $n:[T]^+$. In case of an incoming get, the transmitted value $v$ is remembered as part of $\Delta$ (cf. Equation (8)).

The previous definition deals with the change of context information by communication. Apart from that, unfulfilled promises of the form $n(\bullet)$ also change side, if their name is exchanged together with write-permission. As notation, we will use $C(\Sigma)$ to denote the component $n_1(\bullet) \parallel \cdots \parallel n_k(\bullet)$, where the names $n_i$ correspond to all names of the context $\Sigma$ mentioned as $n_i : [T_i]^+$ for some type $T_i$.

Now to the interface behavior, given by the external steps of Table 10. Most rules have some premises in common. In all cases of a labeled transition, the context $\Sigma$ is updated to $\tilde{\Sigma} = \Sigma + a$ using Definition 2.13. The rules for incoming communication differ from the corresponding ones for outgoing communication in that well-typedness and well-formedness of the label is checked by the premises $\tilde{\Sigma} \vdash [a] : T \rightarrow \bot$, resp. $\tilde{\Sigma} \vdash [a] : _\rightarrow T$ (for calls) resp., $\tilde{\Sigma} \vdash [a] : _\rightarrow T$ (for get-labels), using Definition 2.12. For outgoing communication, the check is unnecessary as starting with a well-typed component, there is no need in re-checking now, as the operational steps preserve well-typedness (subject reduction).

When the component claims the value of a future, we distinguish two situations: the future value is accessed for the first time across the interface or not. In the first case (rules CLAIM1, and CLAIM2), the interface does not contain the value of the future yet, stipulated by the premise $\Delta \vdash n' = \bot$. In that situation it is unclear from the perspective of the component, whether or not the value has already been computed. Hence, it is possible that executing claim is immediately successful (cf. rule CLAIM1) or that the thread $n$ trying to obtain the value has to suspend itself and try later (cf. rule CLAIM2). Rule CLAIM2 works exactly like the corresponding internal rule CLAIMM from Table 5, except that here it is required that the queried future $n'$ is part of the environment. If the future value is already known at the interface (cf. rule CLAIM3 and especially premise $\Delta \vdash n' = v$), executing claim is always successful and the value $v$ is received. Rule CLAIM2 works analogously to claim, except that get insists on obtaining the value, i.e., the alternative of relinquishing the lock and trying again as in rule CLAIM2, is not available for get. The last two rules deal with the situation that the environment fetches the value.

Finally, we characterize the **initial** configuration. Initially, the component contains at most one initial activity and no objects. More precisely, given that $\Sigma_0 \vdash C_0$ is the initial judgment, then $C_0$ contains no objects. Concerning the threads: initially exactly one thread is executing, either at the component side or at the environment side. The distinction is made at the interface that initially either $\Theta_0 \vdash n$ or $\Delta_0 \vdash n$, where $n$ is the only thread name in the system.

### 3. Interface behavior

Next we characterize the possible (“legal”) **interface behavior** as interaction traces between component and environment. Half of the work has been done already in the definition of the external steps in Table 10: for incoming communication, for which the environment is responsible, the assumption contexts are consulted to check whether the communication originates from a realizable environment. Concerning the reaction of the component, no such checks were necessary. To characterize when a given trace is legal, the behavior of the component side, i.e., the outgoing communication, must adhere to the dual discipline we imposed on the environment for the open semantics. This means, we analogously abstract away from the program code, rendering the situation symmetric.

#### 3.1. Legal traces system

The rules of Table 11 specify legality of traces. We use the same conventions and notations as for the operational semantics (cf. Notation 2.11). The judgments in the derivation system are of the form

$$\Sigma \vdash s : \text{trace}.$$  \hspace{1cm} (10)

We write $\Sigma \vdash t : \text{trace}$, if there exists a derivation according to the rules of Table 11 with an instance of L-EMPTY as axiom. The empty trace is always legal (cf. rule L-EMPTY), and distinguishing according to the first action $a$ of the trace, the rules check whether $a$ is possible. Furthermore, the contexts are updated appropriately, and the rules recur checking the tail of the trace. The rules are symmetric wrt. incoming and outgoing communication (the dual rules are omitted). Rule L-CALLI for incoming calls works completely analogously to the CALLI-rule in the semantics: the second premise updates the context $\Sigma$ appropriately with the information contained in $a$, premise $\Sigma' \vdash n$ of L-CALLI assures that the identity $n$ of the future, carrying out the call, is fresh and the two premises $\tilde{\Sigma} \vdash o.l? : T \rightarrow \bot$ and $\tilde{\Sigma} \vdash [a] : T \rightarrow \bot$ together assure that the transmitted values are well-typed (cf. Definition 2.12); the latter two checks correspond to the analogous premises for the external semantics.
in rule CALLI, except that the return type of the method does not play a role here. The L-GetI-rules for claiming a value work similarly. In particular the type checking of the transmitted value is done by the combination of the premises $\Delta \vdash n : [T]$ and $\Xi \vdash [a] : \_ \rightarrow T$. As in the external semantics, we distinguish two cases, namely whether the value of the future has been incorporated in the future already or not (rules L-GetI2 and L-GetI1). In both cases, the thread must be executing on the side of the environment for an incoming get. This is checked by the premise $\Delta \vdash n = \bot$ resp. by $\Delta \vdash n = v$. In case of L-GetI2, where the value of the future has been incorporated as $v$ into the interface information, the actual parameter of the get-label must, of course, be $v$. If not (for L-GetI1), the transmitted argument value is arbitrary, apart from the fact that it must be consistent with the static typing requirements.

It remains to show that the behavioral description, as given by Table 11, actually does what it claims to do, to characterize the possible interface behavior of well-typed components. We show the soundness of this abstraction plus the necessary ancillary lemmas such as subject reduction. Subject reduction means, preservation of well-typedness under reduction. In the formulation of subject reduction, we make sure that the write-permissions of the environment are not available for type-checking the component. We use $[\Delta]$ instead of $\Delta$ as assumption, were $[\_]$ replaces each binding $n$ in $\tau^+$ in $\Delta$ by $n\tau^+$.

**Lemma 3.1** (Subject reduction). If $[\Delta] \vdash C : \Theta$ and $\Delta \vdash C : \Theta \Rightarrow [\Delta] \vdash C : \Theta$, then $[\Delta] \vdash C : \Theta$.

**Proof.** By induction on the number of reduction steps. That internal steps preserve well-typedness, i.e., $[\Delta] \vdash C : \Theta \Rightarrow [\Delta] \vdash C : \Theta$, follows from the corresponding Lemma 2.9 for internal steps. That leaves the external reduction steps of Table 10.

**Case: CALLI**

We are given $[\Delta] \vdash C : \Theta$. The disjunctive premise of the rule distinguishes two sub-cases: 1) $\Sigma' \vdash n$ (where $\Sigma'$ are the bindings carried along with the call-label, i.e., the thread name is transmitted freshly) or 2) $\Delta \vdash n : [\_]^+$ (the thread is not transmitted freshly and the environment has write-permission before the step). Both are treated uniformly in the following argument. For the right-hand side of the transition, we need to show $[\Delta] \vdash C' : [n(letx : T) = \text{grab}(o); M.io(o) in \text{release}(o); x) : \Theta$, where $C'$ corresponds to $C$ extended by new $n'(\_)$-promises. According to the definition of context update (Definition 2.13), $\Xi = \Delta, \Theta$, where $\Theta = \Theta, \Sigma'$, $n[T]^+$ and where $\Sigma'$ contains bindings $n'[T]^+$ for those references transmitted with read-write permission as argument of the call (see the right-hand of equation (8)). The assumption context $\Delta$ for $C$ after the step (by

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**Table 10**

**External steps.**

<table>
<thead>
<tr>
<th>Step</th>
<th>Description</th>
</tr>
</thead>
<tbody>
<tr>
<td>$a = v(\Sigma'). n(call \ a(o)())? \Xi = \Xi+n \Delta \vdash n : [_]^+$ $\Xi \vdash \ a.o : \hat{T} \rightarrow T \Xi \vdash [a] : \hat{T} \rightarrow -$</td>
<td><strong>CALLI</strong></td>
</tr>
<tr>
<td>$\Xi \vdash C \Delta$. $\Xi \vdash C</td>
<td></td>
</tr>
<tr>
<td>$a = v(\Sigma'). \ n(call \ a(o)())? \Xi' = \text{fn}(\ a\ )/ \Xi_1 \ \Xi_1 = \Xi \ \ \ \ \Delta \vdash \ a \vdash \Xi = \Xi + a$</td>
<td><strong>CALLO</strong></td>
</tr>
<tr>
<td>$\Xi \vdash v(\Sigma_1). C(\ C(\Sigma') \ n(letx : T) = \text{bind}(a.o) \rightarrow \text{in} \ t)) \ a \vdash v(\Sigma_1). C(\ C(\Sigma') \ n(letx : T = \text{in} \ t))$</td>
<td></td>
</tr>
<tr>
<td>$a = v(\Sigma'). n'(get(v))? \Xi = \Xi + a \ \ \ \Delta \vdash n' = \perp \ \Xi \vdash [a] : _ \rightarrow T$</td>
<td><strong>CLAIM1</strong></td>
</tr>
<tr>
<td>$\Xi \vdash v(\Sigma_1). C(\ n(letx : T = \text{claim}(\ n', \ a) \ in \ t)) \ a \vdash v(\Sigma_1). C(\ C(\Sigma') \ n(letx : T = \text{in} \ t))$</td>
<td><strong>CLAIM2</strong></td>
</tr>
<tr>
<td>$\Delta \vdash n' = \perp$</td>
<td><strong>CLAIM3</strong></td>
</tr>
<tr>
<td>$\Xi \vdash v(\Sigma_1). C(\ n(letx : T = \text{claim}(\ n', \ a) \ in \ t)) \ a \vdash v(\Sigma_1). C(\ C(\Sigma') \ n(letx : T = \text{release}(o); \text{get}(\ n' \ \text{in} \ \text{get}(o); \ t))$</td>
<td></td>
</tr>
<tr>
<td>$a = n'(get(v))? \Delta \vdash n' = v \Xi \vdash [a] : _ \rightarrow T$</td>
<td><strong>GetI1</strong></td>
</tr>
<tr>
<td>$\Xi \vdash v(\Sigma_1). C(\ n(letx : T) = \text{get}(\ n' \ \text{in} \ t)) \ a \vdash v(\Sigma_1). C(\ C(\Sigma') \ n(letx : T = \text{in} \ t))$</td>
<td><strong>GetI2</strong></td>
</tr>
<tr>
<td>$a = n'(get(v))? \Delta \vdash n' = v \Xi \vdash [a] : _ \rightarrow T$</td>
<td><strong>GetO1</strong></td>
</tr>
<tr>
<td>$\Xi \vdash v(\Sigma_1). C(\ C(\Sigma') \ n(v)) \ a \vdash v(\Sigma_1). C(\ C(\Sigma') \ n(v)) \ a = n'(get(v)!)$</td>
<td><strong>GetO2</strong></td>
</tr>
<tr>
<td>$\Xi \vdash C \ a \vdash C$</td>
<td><strong>GetO2</strong></td>
</tr>
</tbody>
</table>
Table 11
Legal traces (dual rules omitted).

<table>
<thead>
<tr>
<th></th>
<th><img src="https://via.placeholder.com/150" alt="Image" /></th>
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<tbody>
<tr>
<td><img src="https://via.placeholder.com/150" alt="Image" /></td>
<td><img src="https://via.placeholder.com/150" alt="Image" /></td>
</tr>
</tbody>
</table>

the left-hand side of the same equation) is of the form \((\Delta, \Xi') \setminus \Sigma'' \setminus n;[T]^+\). So for the new thread \(n\) at component side, we need to show that

\[
((\Delta, \Xi') \setminus \Sigma'' \setminus n;[T]^+) \vdash [T] : \Theta, \Sigma'', n;[T]^+
\]

with \(t'\) given as \(let_x:T = \text{grab}(o); M.l(o)(\bar{v}) \in \text{release}(o); x\). To derive (11), using a number of instances of \(T-\text{Par}\) in the last derivation steps, gives

\[
\begin{align*}
\Delta, [\Theta], \Sigma'', n;[T]^+ & \vdash t': T \\
\Delta, [\Theta], \Sigma'', n;[T]^+ & \vdash \{ \Delta, [\Theta], \Sigma'', n;[T]^+ \}
\end{align*}
\]

where \(\Delta\) abbreviates the assumption context \((\Delta, \Xi') \setminus \Sigma'' \setminus n;[T]^+\) from (11). Note, how the write permissions from \(\Sigma''\) in commitment of the conclusion at the bottom are split among the three subgoals. All write-permissions are given to the

\[
\begin{align*}
\Xi & \vdash a: s : \text{trace} \\
\Xi & \vdash a: s : \text{trace} \\
\Xi & \vdash a: s : \text{trace}
\end{align*}
\]

the axiom \(T-\text{VAR}\). Remains the right-upper subgoal \((\Delta, \Xi') \setminus \Sigma'' \setminus n;[T]^+\), \([\Theta], \Sigma'', n;[T]^+ \vdash let_x:T = t': T \) (with \(\Delta\) expanded). Note that, apart from the write-permissions, the complicated type context corresponds to \(\Delta, \Xi', \Theta\), or more formally

\[
((\Delta, \Xi') \setminus \Sigma'' \setminus n;[T]^+) \vdash [\Theta, \Sigma'', n;[T]^+]) = [(\Delta, \Xi', \Theta)]
\]

Intuitively, it means, \(t'\) must be checked with all name bindings available from \(\Delta\) and \(\Theta\) plus the ones, which scope is exchanged in \(\Xi'\) as part of the label. No write-permissions, however, can be used to type-check \(t'\) except those being transmitted by the argument of the call and which are kept in \(\Sigma''\) (the context \(\Sigma''\) is the only part of \(t'\) typing context not being stripped off the write-permissions by \(\lfloor\).)

Note that the meta-mathematical notation \(M.l(o)(\bar{v})\) in \(t'\) stands for \(t_{\text{body}}(o/\bar{v})/\bar{x}\), i.e., the method body with the self-parameter \(s\) substituted by the callee’s identity and with the formal parameter replaced by the actual ones. Now, the well-typedness of the pre-configuration \([\Delta] \vdash C: \Theta\) together with the premise \(\Xi \vdash o.l: T \rightarrow T\) of \(\text{CallO}\) (cf. Definition 2.12) implies that \(C\) is of the form \(C : [\xi_1, \ldots, \xi_l = \xi(s,c).\lambda(x:T).t_{\text{body}}(\ldots)]\), and further that \([\Delta] \vdash C: \Theta\) has \(\bar{x}:\bar{T}; [\Delta], \Theta \vdash t_{\text{body}} : T\) as sub-goal. From that, the remaining mentioned subgoal of derivation (12) follows by \(T-\text{LET}\), \(T-\text{GRAB}\), preservation of typing under substitution, \(T-\text{REPLACE}\), and the axiom \(T-\text{VAR}\).

Case: \(\text{CALLO}\)

We are given \(\Delta \vdash v(\Xi_1) : (C \| n;[t]\| n' \langle let_x:T = \text{bind} o.l(\bar{v}) : T \leftarrow n \text{ in } t \rangle) : \Theta\) before the step and \(\Delta \vdash v(\Xi_1) : (C \| n' \langle let_x:T = n \text{ in } t \rangle) : \Theta\) afterwards, with \(C = C \| C_1 \langle \bullet \rangle \| \cdots \| C_n(\bullet)\). By one of the premises of rule \(\text{CALLO}\) we know \(\Delta \vdash o\), i.e., object \(o\) is an environment object.\(^5\) That the name \(o\) refers to an object is assured by the type system and the assumption that the pre-configuration is well-typed. By inverting the rules \(T-\text{Nu}, T-\text{Par}, T-\text{NTHREAD}, T-\text{LET},\) and \(T-\text{BIND}\), we get:

\(^5\) We do not allow cross-border instantiation in this paper, i.e., the component is not allowed to instantiate environment objects and vice versa.
\[ \cdots \frac{Z' = [\Delta], \Xi_1, \Theta \quad Z' = \hat{Z}'(\hat{v}, \hat{t}, \nu[T]^+)}{\text{T-Bind}} \quad \frac{[\Delta], \Xi_1, \Theta \vdash \text{bind} o \, \lambda \hat{v} : T \vdash n : T \quad \xi T ; Z' \vdash t ; T \vdash \cdots}{[\Delta], \Xi_1, \Theta, n' \vdash [T]^+ + \text{let} \, x : T = \text{bind} o \, \lambda \hat{v} : T \vdash n \, \text{in} \, t \vdash [n'[T]^+]++ \quad \text{T-Par, T-Nu} \quad \cdots} \]

Note that \( t \) in the left-upper leaf is type-checked in the context \( Z' \), which corresponds to \( Z' = [\Delta], \Xi_1, \Theta, n'[T]^+ \) with those write-permissions removed that are transmitted via the arguments of the method \( I \) (cf. rule T-Bind).

To derive well-typedness of the post-configuration, we distinguish two sub-cases, namely whether 1) the promise \( n \) is known at the interface before the step or 2) it is hidden still. In the first case, we have \( \Theta \vdash n : T' \) with \( T' = (T)^{++} \) (as a consequence of the fact that the configuration is well-typed), or more precisely, \( \Theta = \hat{\Theta}', \nu[T]^+, n'[T]^+ \). The derivation of well-typedness of the post-configuration \( [\hat{\Delta}] \vdash n'([\xi T] \vdash n \, \text{in} \, t) : \hat{\Theta} \) works as follows:

\[ \frac{[\hat{\Delta}], \hat{\Xi}_1, \hat{\Theta} \vdash n : T \quad \xi T ; [\hat{\Delta}], \hat{\Xi}_1, \hat{\Theta} \vdash t : T'}{\text{T-Let}} \quad \frac{[\hat{\Delta}], \hat{\Xi}_1, \hat{\Theta} \vdash \text{let} \, x : T = n \, \text{in} \, t \vdash n'([T]^+)++ \quad \text{T-NThread} \quad \text{T-Par, T-Nu} \cdots} \]

\[ [\hat{\Delta}] \vdash \nu(\hat{\Xi}_1)(C' \vdash n'([\text{let} \, x : T = n \, \text{in} \, t]) : \hat{\Theta}) \]

where \( \hat{\Theta} = \hat{\Theta}', n'[T]^+ \). Using the premises of the reduction rule CALLO, that relates the different binding contexts mentioned in the step (i.e., \( \hat{\Xi}_1 = \Xi_1 \setminus Z' \), where \( Z' \) are the bindings mentioned in the call labels), \( \hat{\Delta}' = \Delta, \Xi', \nu[T]^+ \) (as stipulated by rule CALLO’s premise \( Z = Z + a \) and given by Definition 2.13, especially equation (8)). Note that (8) is formulated for incoming communication, i.e., used dually here, and that in the considered subcase, we assume that \( n \) is known at the interface before the step, i.e., \( \Theta \vdash n[T]^+ \), as agreed upon earlier. It is straightforward to see that the combined context \( \Delta, \Xi_1, \Theta \) equals \( \Theta, \hat{\Xi}_1, \hat{\Theta} \) with the exception, that the former contains \( n[T]^+ \) (as part of \( \Theta \)) and the latter only \( n'[T]^+ \) (as part of \( \Delta \). Cf. especially by (8)). Furthermore, considering \( [\Delta] \) and \( [\hat{\Delta}] \) in place of \( \Delta \) and \( \Delta \):

\[ [\hat{\Delta}], \hat{\Xi}_1, \hat{\Theta} = ([\Delta], \Xi_1, \Theta) \setminus (n'[T]^+ \cup \hat{\Xi}_1) \]

where \( \hat{v}, \hat{t} \) is given as mentioned in the left-upper leaf if the first derivation tree and as defined by the premise of T-BIND (these bindings correspond to the \( \Sigma \) used in equation (8) and represent the write-permissions transmitted by the call-label from the component to the environment). This discharges the top-left subgoal of the derivation. The second sub-case with \( \Xi_1 \vdash n[T]^+ \) works analogously.

Case: CLAIM\(_1\)

The core of the type preservation here is to assure that the claim-statement in the pre-configuration and the transmitted value \( \nu \) in the post-configuration are of the same appropriate type \( T \). Well-typedness of the pre-configuration implies with claim@\((n', a)\) of type \( T \), that the reference \( n' \) is of type \( [T]^+ \). The third premise of CLAIM\(_1\) states \( Z \vdash [a] : _\rightarrow T \), which implies with Definition 2.12, especially rule LT-GetI of Table 9, that also \( \nu \) is of type \( T \), as required.

Case: CLAIM\(_2\)

By inverting the type rules T-Nu, T-Par, T-Let and T-Claim for the pre-configuration of the step, and by using the same typing rules (except T-Claim) plus T-Get, T-Release, and T-Grab.

The remaining rules work similarly.

**Lemma 3.2** (Soundness of abstractions). If \( \Xi_0 \vdash C \) and \( \Xi_0 \vdash C \Rightarrow \Xi_0 \vdash C \Rightarrow \), then \( \Xi_0 \vdash t \vdash \text{trace} \).

**Proof.** By induction on the number of steps in \( \Rightarrow \). The base case of zero steps (which implies \( t = e \)) is immediate, using L-EMPTY. The induction for internal steps of the form \( \Xi \vdash C \Rightarrow \Xi \vdash C \) follow by subject reduction for internal steps from Lemma 2.9; in particular, internal steps do not change the context \( \Xi \). Remain the external steps of Table 10. First note the contexts \( \Xi \) are updated by each external step to \( \Xi \) the same way as the contexts are updated in the legal trace system.

The cases for incoming communication are checked straightforwardly, as the operational rules check incoming communication for legality, already, i.e., the premises of the operational rules have their counterparts in the rules for legal traces.
Case: \textsc{call}l
Immediate, as the premises of \textsc{l-call}l coincide with the ones of \textsc{call}l.

Case: \textsc{claim}1 and \textsc{get}1
The two cases are covered by rule \textsc{get}1, which has the same premises as the operational rules.

Case: \textsc{claim}2
Trivial, as the step is an internal one.

Case: \textsc{claim}3 and \textsc{get}2
The two cases are covered by \textsc{get}2.

The cases for outgoing communication are slightly more complex, as the label in the operational rule is not type-checked or checked for well-formedness as for incoming communication and as is done in the rules for legality.

Case: \textsc{call}lO
We need to check whether the premises of \textsc{l-call}lO, the dual to \textsc{l-call}l of Table 11, are satisfied. By assumption, the pre-configuration

\[
\Xi \vdash \nu(\Xi_1). (C \parallel n \ (\text{let} \ x:T = \text{bind} \ a.l(\bar{v}) : T \leftarrow n \ \text{in} \ t))
\] (15)

is well-typed. For thread name \(n\) this implies, it is bound either in \(\Xi\) or in \(\Xi_1\), more precisely, either \(\Theta \vdash n : [T]^+(\text{it is public interface information that the component has write-permission for } n)\) or \(\Xi_1 \vdash n : [T]^+\) (the name \(n\) is not yet known in the environment before the communication). In the latter situation we obtain \(\Xi' \vdash n : [\_]^+\) by the premise \(\Xi' = \text{fn}([a]) \cap \Xi_1\) of \textsc{call}lO. Thus, the third premise \(\Xi' \vdash n \lor \Theta \vdash n : [\_]^+\) of \textsc{l-call}lO is satisfied. We furthermore need to check whether the label is type-correct (checked by premises no. 4 and 5 or \textsc{l-call}lO). Its easy to check that the label is well-formed (cf. the first part of Definition 2.12). The first premise of the check of equation (6), that the receiving object \(o\) is an environment object, is directly given by the premise \(\Delta \vdash o\) of \textsc{call}lO. That the object \(o\) supports a method labeled \(l\) (of type \(\bar{T} \rightarrow \bar{T}\)) follows from the fact that the pre-configuration of the call-step is well-typed. So this gives \textsc{l-call}lO's premise \(\Xi \vdash o.l : \bar{T} \rightarrow \bar{T}\). Remains the type check \(\Xi \vdash \bar{v} : \bar{T} \rightarrow \bar{T}\) (checking that the transmitted values \(\bar{v}\) are of the expected type \(\bar{T}\)), which again follows from well-typedness of Eq. (15) (especially inverting \textsc{T-bind}).

The remaining cases work similarly. □

Remark 3.3 (Comparison with reentrant threading). In a multi-threaded setting with synchronous method calls (see for instance [4] [80]), the definition of legal traces is more complicated. Especially, to judge whether a trace is possible requires referring to the past. I.e., instead of judgments of the form of Eq. (10), the check for legality with synchronous calls uses judgments of the form:

\[
\Xi \vdash r > s : \text{trace},
\]

reading "after history \(r\) (and in the context \(\Xi\)), the trace \(s\) is possible". This difference has once more to do with reentrance, resp. with the absence of this phenomenon here. In the thread case, where, e.g., an outgoing call can be followed by a subsequent incoming call as a "call-back". To check therefore, whether a call or a return is possible as a next step involves checking the proper nesting of the call- and return labels. This nesting requirement (also called the balance condition) degenerates here in the absence of call-backs to the given requirement that each call uses a fresh (future) identity and that each get-label (taking the role of the return label in the multithreaded setting) is preceded by exactly one matching preceding call. This can be judged by \(\Delta \vdash n : \_[\_]\) or \(\Theta \vdash n : \_[\_]\) (depending on whether we are dealing with incoming or outgoing get-labels) and especially, no reference to the history of interface interactions is needed. □

Remark 3.4 (Monitors). The objects of the calculus act as monitors as they allow only one activity at a time inside the object. For the operational semantics of Section 2.3, the lock-taking is part of the internal steps. In other words, the handing-over of the call at the interface and the actual entry into the synchronized method body is non-atomic, and at the interface, objects are \textit{input-enabled}. This formalization therefore resembles the one used for the interface description of Java-like reentrant monitors in [3]. To treat the interface interaction and actual lock-grabbing as non-atomic leads to a clean separation of concerns of the component and of the environment. In [3], this non-atomicity, however, gives rise to quite complex conditions characterizing the legal interface behavior. In short, in the setting of [3], it is non-trivial to characterize exactly those situations, when the lock of the object is necessarily taken by one thread which makes certain interactions of other threads impossible. This characterization is non-trivial especially as the interface interaction is non-atomic.

Note, however, that these complications are \textit{not present} in the current setting with active objects, even if the objects act as monitors like in [3]. The reason is simple: there is no need to capture situations when the lock is taken. In Java, the synchronization behavior of a method is part of the interface information. Concretely, the \textit{synchronized}-modifier of Java, specifies that the method’s body is executed atomically in that object without interference of other\footnote{Note that a thread can “interfere” in that setting with itself due to recursion and reentrance.} threads, assuming that \(\text{lock} \rightarrow \bar{T} \rightarrow \bar{T}\).

6 Note that a thread can “interfere” in that setting with itself due to recursion and reentrance.
all other methods of the callee are synchronized, as well. Here, in contrast, there is no interface information that guarantees that a method body is executed atomically. In particular, the method body can give up the lock temporarily via the suspend statement, but this fact is not reflected in the interface information here. This absence of knowledge simplifies the interface description considerably. □

4. Conclusion

We presented an open semantics describing the interface behavior of components in a concurrent object-oriented language with futures and promises. The calculus corresponds to the core of the Creol language, including classes, asynchronous method calls, the synchronization mechanism, and futures, and extended by promises. Concentrating on the black-box interface behavior, however, the interface semantics is, to a certain extent, independent of the concrete language and is characteristic for the mentioned features; for instance, extending Java with futures (see also the citations below) would lead to a quite similar formalization (of course, low level details may be different). Concentrating on the concurrency model, certain aspects of Creol have been omitted here, most notably inheritance and safe asynchronous class upgrades.

Related work

The general concept of “delayed reference” to a result of a computation to be yet completed is quite old. The notion of futures was introduced by Baker and Hewitt [21], where (future e) denotes an expression executed in a separate thread, i.e., concurrently with the rest of the program. As the result of the e is not immediately available, a future variable (or future) is introduced as placeholder, which will eventually contain the result of e. In the meantime, the future can be passed around, and when it is accessed for reading (“touched” or “claimed”), the execution suspends until the future value is available, namely when e is evaluated. The principle has also been called wait-by-necessity [26,27]. Futures provide, at least in a purely functional setting, an elegant means to introduce concurrency and transparent synchronization simply by accessing the futures. They have been employed for the parallel Multilisp programming language [52]. Indeed, quite a number of calculi and programming languages have been equipped with concurrency using future-like mechanisms and asynchronous method calls. Flanagan and Felleisen [42–44] present an operational semantics, based on evaluation contexts, for a λ-calculus with futures. The formalization is used for analysis and optimization to eliminate superfluous dereferencing (“touches”) of future variables. The analysis is an application of a set-based analysis and the resulting transformation is known as touch optimization. Moreau [70] presents a semantics of Scheme equipped with futures and control operators. Promises is a mechanism quite similar to futures and actually the two notions are sometimes used synonymously. They have been proposed in [67]. A language featuring both futures and promises as separate concepts, is Alice ML [17,64,79].

Niehren et al. [72] presents a concurrent call-by-value λ-calculus with reference cells (i.e., a non-purely functional calculus with an imperative part and a heap) and with futures (λ_fut), which serves as the core of Alice ML [17,77,64]. Certain aspects of that work are quite close to the material presented here. In particular, we were inspired by using a type system to avoid fulfilling a promise twice (in [72] called handle error). There are some notable differences, as well. The calculus incorporates futures and promises into a λ-calculus, such that functions can be executed in parallel. In contrast, the notion of futures here, in an object-oriented setting, is coupled to the asynchronous execution of methods. Furthermore, the object-oriented setting here, inspired by Creol, is more high-level. In contrast, λ_fut relies on an atomic test-and-set operation when accessing the heap to avoid atomicity problems. Besides that, they formalize promises using the notion of handled futures, i.e., the two roles of a promise, the writing- and the reading part, are represented by two different references, where the handle to the futures represents the writing-end. Apart from that, [72] are not concerned with giving an open semantics as here. On the other hand, the paper investigates the role of the heap and the reference cells, and gives a formal proof that the only source of non-determinism by race conditions in their language actually are the reference cells and without those, the language becomes (uniformly) confluent.7 Recently, an observational semantics for the (untyped) λ_fut-calculus has been developed in [71]. The observational equivalence is based on may- and must-program equivalence, i.e., two program fragments are considered equivalent, if, for all observing environments, they exhibit the same necessary and potential convergence behavior.

Futures have also been investigated in the object-oriented paradigm. For instance, the object-oriented language Scala [73] has recently been extended [51] by actor-based concurrency, offering futures and promises as part of the standard library. The futures and promises are inspired by their use in Alice ML. In Java 5, futures have been introduced as part of the java.util.concurrent package. As Java does not support futures as a core mechanism for parallelism, they are introduced in a library. Dereferencing of a future is done explicitly via a get-method (similarly to this paper). A recent paper [86] introduces safe futures for Java. The safe concept is intended to make futures and the related parallelism transparent and in this sense goes back to the origins of the concept: introducing parallelism via futures does not change the program’s meaning.

7 Uniform confluence is a strengthening of the more well-known notion of (just ordinary) confluence; it corresponds to the diamond property of the one-step reduction property. For standard reduction strategies of a purely functional λ-calculus, only confluence holds, but not uniform confluence. However, the non-trivial “diamonds” in the operational semantics of λ_fut are caused not by different redexes within one λ-term (representing one thread), but by redexes from different threads running in parallel, where the reduction strategy per thread is deterministic (as in our setting, as well).
straightforward and natural in a functional setting, safe futures in an object-oriented and thus state-based language such as 
Java require more considerations. The paper introduces a semantics which guarantees safe, i.e., transparent, futures by 
deriving restrictions on the scheduling of parallel executions and uses object versioning. The futures are introduced as an 
extension of Featherweight 
Java (\(F\)) [54], a core object calculus, and is implemented on top of 
Jikes RVM [18,25]. Pratikakis et al. [75] present a constraint-based static analysis for (transparent) futures and proxies in 
Java, based on type qualifiers and qualifier inference [45]. Also this analysis is formulated as an extension of \(F\) by type qualifiers. Similarly, Carmel et al. [28–30] tackle the problem to provide \textit{confluent}, i.e., effectively deterministic system behavior for a concurrent object 
calculus with futures (asynchronous sequential processes, \textit{ASP}, an extension of Abadi and Cardelli’s imperative, untyped 
object calculus \textit{imp}_c [1]) and in the presence of imperative features. The \textit{ASP} model is implemented in the \textit{ProActive Java}- 
library [31]. The fact, that \textit{ASP} is derived from some (sequential, imperative) object-calculus, as in the formalization here, 
is more a superficial or formal similarity, in particular when being interested in the interface behavior of concurrently 
running objects, where the inner workings are hidden anyway. Apart from that there are some similarities and a number of 
differences between the work presented here and \textit{ASP}. First of all, both calculi are centered around the notion of first-class 
futures, yielding active objects. The treatment, however, of getting the value back, is done differently in [28]. Whereas here, 
the client must explicitly claim a return value of an asynchronous method, if interested in the result, the treatment of the 
future references is done \textit{implicitly} in \textit{ASP}, i.e., the client blocks if he performs a strict operation on the future (without explicit 
syntax to claim the value). Apart from that, the object model is more sophisticated, in that the calculus distinguishes between 
active and passive objects. Here, we simple have objects, which can behave actively or passively (reactively), depending on 
the way they are used. In \textit{ASP}, the units of concurrency are the explicitly activated active objects, and each passive one is 
owned and belongs to exactly one active one. Especially, passive objects do not directly communicate with each other across 
the boundaries of concurrent activity, all such communication between concurrent activities is mediated and served by the 
active objects.

Related to that, a core feature of \textit{ASP}, not present here, is the necessity to specify (also) the \textit{receptive behavior} of the 
active object, i.e., in which order it is willing to process or serve incoming messages. The simplest serve strategy would be 
the willingness to accept all messages and treat them in a first-come, first-serve manner, i.e., a input-enabled FIFO strategy 
on the input message queue. The so-called \textit{serve}-method is the dedicated activity of an active object to accept and schedule 
incoming method calls. Typically, as for instance in the FIFO case, it is given an an non-terminating process, but it might 
also terminate, in which case the active object together with the passive objects it governs, becomes superfluous: an active 
object which does no service any longer does not become a passive data structure, but can no longer react in any way.

As extension of the core \textit{ASP} calculus, [28, Chapter 10] treats \textit{delegation} that bears some similarities with the promises 
here. By executing the construct \textit{delegate}(\textit{o.l(\vec{v}))} (using our notational conventions), a thread \textit{n} hands over the permission 
and obligation to provide eventually a value for the future reference \textit{n} to method \textit{l} of object \textit{o}, thereby losing that permission 
itself. That corresponds to executing \textit{bind} \textit{o.l(\vec{v})} : \textit{T} \leftarrow \textit{n}. Whereas in our setting, we must use a yet-unfulfilled promise \textit{n} for 
that purpose, the delegation operator in \textit{ASP} just (re-)uses the current future for that. Consequently, \textit{ASP} does not allow the 
creation of promises independently from the implicit creation when asynchronously calling a method, as we do with the 
promise \textit{T} construct. In this sense, the promises here are more general, as they allow to profit from delegation and have the 
promise as first-class entity, i.e., the programmer can pass it around, for instance, as argument of methods. This, on the other 
hand, requires a more elaborate type system to avoid write errors on promises. This kind of error, fulfilling a promise twice, 
is avoided in the delegate-construct of \textit{ASP} not by a type system, but by construction, in that the delegate-construct must be 
used only at the end of a method, so that the delegating activity cannot write to the future/promise after it has delegated 
the permission to another activity.

Further uses of futures for \textit{Java} are reported in [68,62,76,83,82]. Futures are also integral part of \textit{Io} [55] and \textit{Scoop} (simple 
concurrent object-oriented programming) [35,20,69], a concurrent extension of \textit{Eiffel}. Both languages are based on the active 
objects paradigm.

Benton et al. [23] present polyphonic \textit{C^*}, adding concurrency to \textit{C^*}, featuring asynchronous methods and based on the 
join calculus [46,47]. Polyphonic \textit{C^*} allows methods to be declared as being asynchronous using the \textit{async} keyword for the 
method type declaration. Besides that, polyphonic \textit{C^*} supports so-called \textit{chords} as synchronization or join pattern. With 
similar goals, \textit{Java} has been extended by join patterns in [56,57].

In the context of \textit{Creol}, de Boer et al. [39] present a formal, operational semantics for the language and extend it by \textit{futures} 
(but not promises). Besides the fact, that both operational semantics ultimately formalize a comparable set of features, there 
are, at a technical level, a number of differences. For once, here, we simplified the language slightly mainly in two respects 
(apart from making it more expressive in adding promises, of course). We left out the “interleaving” operators \(\parallel\) and \(\&\) of 
[39] which allows the user to express interleaving concurrency \textit{within} one method body. Being interested in the observable 
interface behavior, those operations are a matter of internal, hidden behavior, namely leading to non-deterministic behavior 
at the interface. Since objects react non-deterministically anyhow, namely due to race conditions present independently of 
\(\parallel\) and \(\&\), those operators have no impact on the possible traces at the interface. The operators might be useful as abstractions 
for the programmer, but without relevance for the interface traces, and so we ignored them here. Another simplification, this 
time influencing the interface behavior, is how the programmer can claim the value of a future. This influences, as said, the 
interface behavior, since the component may fetch the value of a future being part of the environment, or vice versa. Now, 
the design of the \textit{Creol}-calculus in [39] is more liberal wrt. what the user is allowed to do with future references. In this paper, 
the interaction is rather restricted: if the client requests the value using the claim-operation, if the client are basically only two
formulated the (internal) operational semantics using evaluation contexts (as do, e.g. [72]) for a future reference observable and would complicate the behavioral interface semantics to some extent. This is also corroborated by the circumstance that the expressive power of explicit polling quite complicates the proof theory of [39] (see also the discussion in the conclusion of [39]). This is not a coincidence, since one crux of the complete Hoare-style proof systems such as in [39] is to internalize the (ideally observable) behavior into the program state by so-called auxiliary variable. In particular recording the past interaction behavior in so-called history variables is, of course, an internalization of the interface behavior, making it visible to the Hoare-assertions. As a further indication that allowing to poll a future quite adds expressivity to the language is the observation that adding a poll-operation to ASP, destroys a central property of ASP, namely confluence, as is discussed in [28, Chapter 11].

Apart from that, the combination of claiming a futures, the possibility of polling a future, and a general await-statement complicates the semantics of claiming a future: in [39], this is done by busy-waiting, which in practice one intends to avoid. So instead of the behavior described in Fig. 1, the formalization in [39] behaves as sketched in Fig. 3.

After an unsuccessful try to obtain a value of future, the requesting thread is suspended and loses the lock. In order to continue executing, the blocked thread needs two resources: the value of the future, once it is there, plus the lock again. The difference of the treatment in Fig. 1 and the one of Fig. 3 for [39] is the order in which the requesting thread attempts to get hold of these two resources: our formalization first check availability of the future and afterwards re-gains the lock to continue, whereas [39] do it vice versa, leading to busy wait. The reason why it is sound to copy the future value into the local state space without already having the lock again (Fig. 1) is, of course, that, once arrive, the future value remains stable and available.

In addition, our work differs also technically in the way, the operational semantics is represented. de Boer et al. [39] formulated the (internal) operational semantics using evaluation contexts (as do, e.g. [72] for \( \lambda_{\text{hoare}} \)), whereas we rely on a “reduction-style” semantics, making use of an appropriate notion of structural congruence. While largely a matter of taste, it seems to us that, especially in the presence of complicated synchronization mechanisms, for instance the ready queue representation of [39], the evaluation contexts do not give rise to an immediately more elegant specification of the reduction behavior. Admittedly, we ignored here the internal interleaving operators \(||\parallel\) and \(\llnot\parallel\), which quite contribute to the complexity of the evaluation contexts. Another technical difference, if you wish, concerns the way, the futures, threads, and objects are represented in the operational semantics, i.e., in the run-time syntax of the calculus. Different from our representation, their semantics makes the active-objects paradigm of Creol more visible: The activities are modeled as part of the object. More precisely, an object contains, besides the instance state, an explicit representation of the current activity (there called “process”) executing “inside” the object plus a representation of the ready-queue containing all the activities, which have been suspended during their execution inside the object. The scheduling between the different activities is then done by juggling them in and out of the ready-queue at the processor release points. Here, in contrast, we base our semantics on a separate representation of the involved semantics concepts: (1) classes as generators of objects, (2) objects carrying in the instance variables the persistent state of the program, thus basically forming the heap, and (3) the parallel activities in the form of threads. While this representation makes arguably the active-object paradigm less visible in the semantics, it on the other hand separates the concepts in a clean way. Instead of an explicit local scheduler inside the objects, the access to the shared instance states of the objects is regulated by a simple, binary lock per object. So, instead of having two levels of parallelism – locally inside the objects and inter-object parallelism – the formalization achieves the same with just one conceptual level, namely: parallelism is between threads (and the necessary synchronization is done via object-locks). Additionally, our semantics is rather close to the object-calculi semantics for multi-threading as in Java (for instance as in [58,59] or [80]). This allows to see the differences and similarities between the different models of concurrency, and the largely similar representation could allow are more formal comparison between the interface behaviors in the two settings.

The language Cool [32,33] (concurrent, object-oriented language) is defined as an extension of C++ [81] for task-level parallelism on shared memory multi-processors. Concurrent execution in Cool is expressed by the invocation of parallel functions executing asynchronously. Unlike the work presented here, Cool contains future types, which correspond to the types of the form \( T \) used here. Further languages supporting futures include ACT-1 [65,66], concurrent Smalltalk [87,91], and of course the influential actor model [16,50,15], ABCL/f [88,89] (in particular the extension ABCL/[84]).
We have characterized the behavioral semantics of open systems, similarly to the one presented here for futures and promises, in earlier papers, especially for object-oriented languages based on Java-like multi-threading and synchronous method calls, as in Java or C#. The work [5] deals with thread classes and [4] with re-entrant monitors. In [80] the proofs of full abstraction for the sequential and multi-threaded cases of a class-based object-calculus can be found. Poetzsch-Heffter and Schäfer [74] present a behavioral interface semantics for a class-based object-oriented calculus, however without concurrency. The language, on the other hand, features an ownership-structured heap.

Future work

An obvious way to proceed is to consider more features of the Creol-language, in particular inheritance and subtyping. Incorporating inheritance is challenging, as it renders the system open wrt. a new form of interaction, namely the environment inheriting behavior from a set of component classes or vice versa. Also Creol’s mechanisms for dynamic class upgrades should be considered from a behavioral point of view (that we expect to be quite more challenging than dealing with inheritance). An observational, black-box description of the system behavior is necessary for the compositional account of the system behavior. Indeed, the legal interface description is only a first, but necessary, step in the direction of a compositional and ultimately fully-abstract semantics, for instance along the lines of [80]. Based on the interaction trace, it will be useful to develop a logic better suited for specifying the desired interface behavior of a component than enumerating allowed traces. Another direction is to use the results in the design of a black-box testing framework, as we started for Java in [38]. We expect that, with the theory at hand, it should be straightforward to adapt the implementation to other frameworks featuring futures, for instance, to the future libraries of Java 5.

Acknowledgements

We thank Einar Broch Johnsen, Marcel Kyas, Olaf Owe, and Gerardo Schneider for stimulating discussions on the topic. Furthermore we are thankful for the detailed comments and careful suggestions of the reviewers that helped improving the paper.

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<td>name (in general, or thread / future / promise name in particular)</td>
</tr>
<tr>
<td>( v )</td>
<td>493</td>
<td>value</td>
</tr>
<tr>
<td>( o )</td>
<td>493</td>
<td>object name</td>
</tr>
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</tr>
<tr>
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<td>494</td>
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<td>( c[x] )</td>
<td>494</td>
<td>class with name ( c ) and methods and fields defined in ( O )</td>
</tr>
<tr>
<td>( o[c,F,L] )</td>
<td>494</td>
<td>instance ( o ) of class ( c ) with fields ( F ) and lock ( L )</td>
</tr>
<tr>
<td>( \top ), ( \bot )</td>
<td>494</td>
<td>value of a taken, resp., free lock</td>
</tr>
<tr>
<td>( \nu )</td>
<td>494</td>
<td>( \nu )-operator for hiding</td>
</tr>
<tr>
<td>( \varsigma (s:T).\lambda (\vec{x}:\vec{T}).t )</td>
<td>494</td>
<td>method of the abstracted object ( s ) with formal parameters ( \vec{x} ) and body ( t )</td>
</tr>
<tr>
<td>( \bot_c )</td>
<td>494</td>
<td>undefined object reference</td>
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<tr>
<td>( f )</td>
<td>494</td>
<td>field name</td>
</tr>
<tr>
<td>( l = \varsigma (s:T).\lambda (\vec{x}:\vec{T}).t ) (or ( l = \nu )) resp. ( l = \varsigma (s:T).\lambda (\vec{x}:\vec{T}).t ) (or ( l = \bot_c ))</td>
<td>494</td>
<td>field definition</td>
</tr>
<tr>
<td>( v.l() ) (or ( v.l ))</td>
<td>494</td>
<td>field access</td>
</tr>
<tr>
<td>( v’.l := \varsigma (s:T).\lambda (\vec{x}:\vec{T}).t ) (or ( v’.l := \nu ))</td>
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<td>field update</td>
</tr>
<tr>
<td>( v_{\bot} )</td>
<td>494</td>
<td>either a value ( v ) or a symbol ( \bot_c )</td>
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<tr>
<td>new ( c )</td>
<td>494</td>
<td>object creation</td>
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<td>promise ( T )</td>
<td>494</td>
<td>promise creation</td>
</tr>
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<td>494</td>
<td>binding code to a promise</td>
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<td>claim, get</td>
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### 2.2 Type system

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<td>$B$</td>
<td>496</td>
<td>base types (such as integers, etc. Left unspecified)</td>
</tr>
<tr>
<td>Unit</td>
<td>496</td>
<td>type of the unit value ()</td>
</tr>
<tr>
<td>$[T]^+$</td>
<td>496</td>
<td>type of a reference to a future with write permission, where the future will return a value of type $T$</td>
</tr>
<tr>
<td>$[T]^+$</td>
<td>496</td>
<td>type of a reference to a future with read-only permission, where the future will return a value of type $T$</td>
</tr>
<tr>
<td>$[T]$</td>
<td>496</td>
<td>type of a reference to a future with write or read-only permission, where the future will return a value of type $T$</td>
</tr>
<tr>
<td>$[l_1;U_1,\ldots,l_k;U_k]$</td>
<td>496</td>
<td>interface type of unnamed objects</td>
</tr>
<tr>
<td>$(l_1;U_1,\ldots,l_k;U_k)$</td>
<td>496</td>
<td>interface type for classes</td>
</tr>
<tr>
<td>$\vec{T}$</td>
<td>496</td>
<td>$T_1 \times \ldots \times T_k$</td>
</tr>
<tr>
<td>Unit $\to T$</td>
<td>496</td>
<td>$T_1 \times \ldots \times T_k \to T$ when $k = 0$</td>
</tr>
<tr>
<td>$\Delta$</td>
<td>496</td>
<td>name context: typing assumptions about the environment</td>
</tr>
<tr>
<td>$\Theta$</td>
<td>496</td>
<td>name context: typing commitments of the component</td>
</tr>
<tr>
<td>$\Delta \vdash C : \Theta$</td>
<td>496</td>
<td>typing judgment</td>
</tr>
<tr>
<td>$\uplus$</td>
<td>496</td>
<td>symmetric operation on well-formed name contexts</td>
</tr>
<tr>
<td>$\Gamma$</td>
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<td>stack-organized variable context</td>
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<td>$n(\epsilon)$</td>
<td>497</td>
<td>not yet fulfilled promise</td>
</tr>
<tr>
<td>$\leq$</td>
<td>497</td>
<td>subtyping relation on types</td>
</tr>
<tr>
<td>$\Gamma ; \Delta \vdash e : T :: \vec{T}, \Delta$</td>
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<td>judgment: type system</td>
</tr>
<tr>
<td>$\Delta \setminus n : T, \Gamma \setminus n : T, \ldots$</td>
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### 2.3 Operational semantics

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<td>internal transition, confluent step</td>
</tr>
<tr>
<td>$\rightarrow$</td>
<td>499</td>
<td>internal transition, other (non-confluent) step</td>
</tr>
<tr>
<td>$\equiv$</td>
<td>500</td>
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<tr>
<td>$\vdash C : ok$</td>
<td>501</td>
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</tr>
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<td>$\Delta ; m C : \Theta$</td>
<td>501</td>
<td>minimal typing judgment</td>
</tr>
<tr>
<td>$\Delta \vdash n = v$</td>
<td>504</td>
<td>$\Delta$ contains the corresponding value information</td>
</tr>
<tr>
<td>$\Delta \vdash n = \bot$</td>
<td>504</td>
<td>$\Delta$ does not contain the corresponding value information</td>
</tr>
<tr>
<td>$\Delta \vdash C : \Theta \xrightarrow{a} \check{\Delta} \vdash \vec{C} : \check{\Theta}$</td>
<td>504</td>
<td>external transition</td>
</tr>
<tr>
<td>$\Xi$</td>
<td>504</td>
<td>the tuple of name contexts $\Delta, \Theta$</td>
</tr>
<tr>
<td>$\gamma c$</td>
<td>504</td>
<td>call label</td>
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<td>$\gamma g$</td>
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</tr>
<tr>
<td>$\gamma ?$</td>
<td>504</td>
<td>incoming interaction label</td>
</tr>
<tr>
<td>$\gamma !$</td>
<td>504</td>
<td>outgoing interaction label</td>
</tr>
<tr>
<td>$\lfloor \gamma \rfloor$</td>
<td>504</td>
<td>core of the label $\gamma$</td>
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<tr>
<td>$fn(a)$</td>
<td>504</td>
<td>free names of label $a$</td>
</tr>
<tr>
<td>$bn(a)$</td>
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<td>bound names of label $a$</td>
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<tr>
<td>$names(a)$</td>
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<td>all names of label $a$</td>
</tr>
<tr>
<td>$fn_a(a)$</td>
<td>504</td>
<td>free names occurring in active position in $a$</td>
</tr>
<tr>
<td>$fn_p(a)$</td>
<td>504</td>
<td>the free names in passive position in $a$</td>
</tr>
<tr>
<td>$\vdash a$</td>
<td>505</td>
<td>$a$ is well-formed</td>
</tr>
<tr>
<td>$\Xi \vdash o.l? : \vec{T} \to T$</td>
<td>505</td>
<td>an incoming call of the method labeled $l$ in object $o$ expects arguments of type $\vec{T}$ and results in a value of type $T$</td>
</tr>
<tr>
<td>$\Xi \vdash a : T \to _\text{resp.}, \check{\Xi} \vdash a : _ \to T$</td>
<td>505</td>
<td>well-typedness of an incoming core label $a$ with expected type $\vec{T}$, resp., $T$, and relative to the name context $\Xi$</td>
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</table>
2.3 Operational semantics

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<tr>
<td>$\Xi \vdash v : T$</td>
<td>505</td>
<td>$\Xi_0$ abbreviate $\Xi$, then $\Xi_i \vdash v_i : T_i$ and $\Xi_{i+1} = \Xi_i \setminus T_i$, for all $0 \leq i \leq n - 1$</td>
</tr>
<tr>
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3 Interface behavior

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