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JSkel: Towards a Formalization of JavaScript’s Semantics

Adam Khayam, Louis Noizet, and Alan Schmitt

Inria

Abstract

We present JSkel, a formalization of the semantics of JavaScript in Skel, the concrete language used to write skeletal semantics. We describe the improvements to Skel we designed and implemented to significantly simplify the formalization. We show the formalization is both close to the specification and executable.

1 Introduction

Mechanizing the semantics of complex languages is useful, not only to make it precise, but also to obtain a framework in which one may prove properties of programs or of the language itself. The success of such a mechanization may be evaluated in two ways: is the actual semantics of the language really captured, and can the formalization be used in practice?

There are two main ways to make sure the semantics is captured: first, the mechanization can be textually close to the specification of the language, if it exists. Second, if the mechanization is executable, it should successfully run tests provided by the language. Thus, a suitable tool to mechanize semantics should make it easy to stay close to language definitions, be they algorithmic (e.g., for JavaScript) or based on inference rules. Furthermore, the tool should also provide an easy way to derive an executable version of the semantics to run code, including tests.

The goal of this paper is to show that the Skeletal Semantics framework [2] is suitable to mechanize semantics. To this end, we describe the ongoing formalization of JavaScript as a skeletal semantics. In this paper, we only address the question of capturing the semantics. We leave the evaluation of the usability of the formalization for future work.

JSkel is the first formalization of a real-world programming language in Skel and it has been a core support in the evolution of Skel, from the introduction of the notion of polymorphism to the addition of monads and first-class functions. The evolution of Skel is motivated by the need of having a formal semantics and a language expressive enough to be able to capture the behavior of JavaScript, ensuring a visual and a behavioral match between the formalization and the specification.

We present existing approaches to mechanize languages in Section 2. We then introduce Skeletal Semantics as well as our concrete syntax in Section 3. Our main contribution is the description of the mechanization of JavaScript in Section 4, where we also present some crucial features of skeletal semantics that significantly improve the readability of the semantics. We conclude and explore future work in Section 5.

2 Context

We first motivate why we chose JavaScript to evaluate the mechanization of a language using skeletal semantics. The three defining features of JavaScript in this regard are the following. First, it is complex, hence a good candidate to see if our solution scales up. Second, it has a precise specification, called ECMAScript (ES), and a large suite of test cases, hence we do not
have to guess what its semantics is. And third, it has been mechanized in several frameworks, which facilitates the comparison to other approaches.

In addition, we have a previous experience in mechanizing JavaScript in Coq, called JSCert [1]. Defining a semantics in a proof assistant is the most direct way of mechanizing it. It has a major drawback, however. If the design choices are not correct, one cannot manipulate the mechanization to modify it, and one must instead redo it using different choices. JSCert is a pretty-big-step [4] Coq definition of ECMAScript 5.1. It consists of an inductive definition, a recursive definition, and a correctness proof showing they match. The goal of the inductive definition is to prove properties of the language and of JavaScript programs, whereas the recursive definition can provide an OCaml interpreter using Coq’s extraction mechanisms. The mechanization is fairly close to the specification for people who can read Coq code, and the test suite can be run using the extracted interpreter. JSCert has two flaws, unfortunately. It is difficult to maintain, as one has to update two formalizations and a proof. In addition, and most importantly, JSCert uses Coq’s induction to represent the recursive evaluation of the language. This shallow embedding results in a definition of a semantics as an inductive of about 1000 rules, which is too large to prove properties of the language. The exploration of ways to address these issues was the motivation to create skeletal semantics.

To simplify the mechanization, [6] introduces a core language $\lambda_{JS}$ that embodies JavaScript essentials, and a desugaring process that transforms a JS program to a $\lambda_{JS}$ one. This work is based on ECMA 3 and 5. The authors empirically prove that the desugared version of JS is compliant with the implementations of JS itself using JavaScript’s test suites. In spite of its efficiency and correctness, $\lambda_{JS}$ is a feature-based redefinition of ECMA. This formalization does not follow the specification, making it hard to gather evidence it actually captures the language beyond running the tests. We attempted to formalize $\lambda_{JS}$ in Coq\(^1\), uncovering several issues with the desugaring process that were not witnessed by testing.

An alternative approach to the formalization of JavaScript is the use of an existing framework. KJS [11] is a complete mechanization of ECMAScript 5.1 in the K framework. It provides an executable interpreter of JS directly from the semantics with no additional effort. This framework is not suitable to analyze the language itself as it only provides tools to reason about the execution of programs. In addition, there is no evidence that the mechanization can be easily maintained: although JavaScript has significantly evolved since ES 5.1 (the specification has more than doubled in size), the mechanization has not been updated. Our experience in updating JSExplain [3], a JavaScript interpreter written in OCaml, from ES5.1 to ES6 has shown us it is far from a trivial issue. The power of the K framework comes at the cost of additional complexity in the description of languages in it, which hinders maintainability and closeness to the specification.

### 3 Skeletal Semantics

Skeletal semantics is a **syntax** to define the semantics of programming languages in a concise yet powerful way, with a light formalism. This provides a way to easily manipulate the semantics, for instance to convert it into a Coq formalization or an OCaml interpreter. One of the strengths of skeletal semantics is the possibility to leave some constructions undefined, to let them be implementation dependent or for gradual specification.

The theoretical concept behind skeletal semantics was presented in [2]. The version we show here has been significantly improved, but it is still work in progress, in particular regarding some

---

\(^1\)https://github.com/tilk/LambdaCert
eval_stmt (σ, While (e, t)) :=
    let b = eval_expr (σ, e) in
    \[
    \begin{cases}
    \text{let True } = b \text{ in } \sigma' = \text{eval_stmt } (\sigma', \text{While } (e, t)) \\
    \text{let False } = b \text{ in } \sigma
    \end{cases}
    \]

Figure 1: skeleton for the While constructor

Theoretical aspects such as the definition of abstract interpretations.

The “Skel” language, which has been defined to describe skeletal semantics, serves as support for the necro ecosystem [9], which provides a generator of OCaml interpreters [7], a generator of Gallina formalization [8], and a (work in progress) generator of \TeX{} inference rules [10].

The Skel formalization of JavaScript, which constitutes the main topic of this article, has had a lot of impact on the language itself, as it prompted us to make several improvements to Skel, such as adding polymorphism and first-class functions.

Figure 1 shows an example of a skeleton for the evaluation of a while block. We see all the main elements of skeletal semantics and we can observe that they are actually the main elements of any semantics.

- Recursion (eval_stmt and eval_expr) lets us define the semantics depending on the semantics of other terms (frequently subterms, but not always, as we can see in this example with the call eval_stmt (σ', While (e, t))).
- There are auxiliary functions and auxiliary types such as booleans, and possibility to match a boolean variable against a given constructors (e.g. True).
- The let...in construct is used to perform a sequence of operations.
- The branching (represented as a parenthesized system in Figure 1) is a choice between two possible rules. Often, the branches are mutually exclusive and a pattern matching at the start of the branch determines which branch is taken. Non-mutually exclusive branches lets us also represent non-deterministic semantics (such as \(\lambda\)-calculus with no evaluation strategy).

3.1 Formalism

We ignore polymorphism in the formal definition below to keep it readable. The use of polymorphism is pretty intuitive though, and its semantics is the usual one. We first give the syntax of skeletal semantics.

\[
\begin{align*}
\text{TERM} & \quad t \quad ::= \quad x_i \mid C \ t \mid (t,\ldots,t) \mid \lambda x : \tau . \ S \\
\text{SKELETON} & \quad S \quad ::= \quad x_i \ t_1 \ldots t_n \mid \text{let } p = S \text{ in } S \mid (S . S) \mid t \\
\text{PATTERN} & \quad p \quad ::= \quad x_i \mid . \mid C \ p \mid (p,\ldots,p) \\
\text{TYPE} & \quad \tau \quad ::= \quad b \mid \tau \rightarrow \tau \mid (\tau,\ldots,\tau)
\end{align*}
\]
A term is either a variable, a constructor applied to a term, a tuple of terms, or an abstraction whose body is a skeleton (where the type of the abstracted variable is explicitly given). Terms can be viewed as expressions in their normal form (that is with no reduction possible). A skeleton is either the application of a variable to several terms, a let binding, where the bound skeleton is matched against a pattern, a branching of several skeletons, or simply a term (sometimes, we write $\text{ret } t$ to say explicitly that we consider the skeleton and not the term). A pattern is either a variable name, a wildcard, a constructor applied to a pattern, or a tuple of patterns. Finally, a type is either a base type, that is a type declared by the user (either specified or unspecified), an arrow type, or a tuple of types.

This syntax is very close to Abstract Normal Forms [5]. The main difference is that we add branching as a non-deterministic choice. We also provide constructors and tuple that are not in the basic ANF, and we allow let bindings in let bindings. Of course, the let bindings can always be extracted by using a fresh variable name, which the function $\text{necrotrans extractletin}$ actually does.

A skeletal semantics is a list of type declarations, either unspecified or specified (by giving its constructors), and a list of term declarations, also either unspecified and containing only a name and type, or specified with an additional term.

We give the following four examples, written in Skel:

\begin{verbatim}
type int

| Zero | Succ nat

type nat = | Zero | Succ nat

term add: int → int → int

term two: nat = Succ (Succ Zero)
\end{verbatim}

The possibility to declare non-specified types and terms is a really powerful tool. When defining a semantics, we sometimes do not want to go into details on how every type and every function works. Or sometimes, we want it to remain unspecified. The partial specifications allows for that.

As shown above, the Skel language is really light (close to $\lambda$-calculus), yet very powerful. A good proof of this power is the semantics of JS given in Section 4.

To furthermore improve readability, the Skel language also provide notation for binding operators. That is, assuming a term declaration for $\text{bind}$, one can write $\text{let}\%\text{bind } p = x \text{ in } v$ for $\text{bind } x \ (\lambda v -> \text{let } p = v \text{ in } s)$. It is particularly useful when writing a semantics that heavily uses monadic constructions, such as the one for JavaScript.

In itself, the language is only a syntactic construct, but there are multiple ways to derive meaning out of a skeletal semantics. This enables the writing of a single skeletal semantics to obtain multiple semantics. The usual way to interpret a skeletal semantics is the concrete semantics, which corresponds to a natural (big step) semantics. We describe it in Section 3.3, but first we focus on typing.

### 3.2 Typing

Skel is strongly typed. As mentioned above, types are threefold: user-defined types (specified or unspecified), product types for tuples, and arrow types for functions.

We give the typing rules for terms and skeletons in Figure 2. They are respectively of the form $\Gamma \vdash_t t : \tau$ and $\Gamma \vdash_S S : \tau$, where $\Gamma$ is a typing environment (a partial functions that binds variable names to types). The initial typing environment $\Gamma_0$ is the function that binds every term name (both specified and non-specified) to the associated type given in the term declaration list.
\[
\begin{align*}
\Gamma(x) &= \tau \\
\Gamma \vdash t : \tau &\quad \text{VAR} \\
\Gamma \vdash C : \tau' &\quad \text{CONST} \\
\Gamma \vdash t_1 : \tau_1 &\quad \ldots \quad \Gamma \vdash t_n : \tau_n \quad \text{TUPLE} \\
\Gamma \vdash (t_1, \ldots, t_n) : (\tau_1, \ldots, \tau_n) &\quad \text{CLOS} \\
\Gamma \vdash \lambda x : \tau \cdot S : \tau' &\quad \text{LETIN} \\
\Gamma \vdash S : \tau &\quad \Gamma \vdash S_1 : \tau &\quad \ldots \quad \Gamma \vdash S_n : \tau &\quad \text{BRANCH} \\
\forall i \quad \Gamma \vdash t_i : \tau_i &\quad \Gamma(x) = \tau_1 \to \cdots \to \tau_n \to \tau &\quad \text{APP} \\
\Gamma \vdash \text{ret} : \tau &\quad \Gamma \vdash \text{let} \; S \; \text{in} \; S' : \tau' &\quad \text{LETIN} \\
\Gamma + x \leftarrow \tau &\quad \text{is the partial function } \Gamma' \text{ s.t. } \Gamma'(x) = \tau \text{ and for all } y \neq x, \Gamma'(y) = \Gamma(y). \\
\Gamma + \_ \leftarrow \tau &= \Gamma. \\
\Gamma + C \; p \leftarrow \tau' &= \Gamma + p \leftarrow \tau \text{ if } \text{ctype}(C) = (\tau, \tau'). \\
\Gamma + (p_1, \ldots, p_n) \leftarrow (\tau_1, \ldots, \tau_n) &= \Gamma + p_1 \leftarrow \tau_1 + \cdots + p_n \leftarrow \tau_n.
\end{align*}
\]

To type a variable, we look for its type in \(\Gamma\). To type a constructor applied to a term, we type the given term, check that it matches the requirement for the constructor, and return the output type of the constructor. Both are given using \(\text{ctype}(C)\) which returns the pair of the declared input type and output type for the constructor \(C\) as found in the type declaration list. To type a tuple, we type each component and return the tuple of the types. Finally, to type an abstraction, we type the skeleton, and return the arrow from the type of the argument to the type of the skeleton.

To type a skeleton that is a term, we simply type the term. To type a branching, we require that every branch has the same type \(\tau\), and we return \(\tau\) (an empty branch therefore can be typed to any given type). To type a let binding, we first type the bound skeleton, then we type the continuation in the extended environment. To this end, we introduce the extensions of \(\Gamma\) with a matched type.

- \(\Gamma + x \leftarrow \tau\) is the partial function \(\Gamma'\) s.t. \(\Gamma'(x) = \tau\) and for all \(y \neq x\), \(\Gamma'(y) = \Gamma(y)\).
- \(\Gamma + \_ \leftarrow \tau = \Gamma\).
- \(\Gamma + C \; p \leftarrow \tau' = \Gamma + p \leftarrow \tau\) if \(\text{ctype}(C) = (\tau, \tau')\).
- \(\Gamma + (p_1, \ldots, p_n) \leftarrow (\tau_1, \ldots, \tau_n) = \Gamma + p_1 \leftarrow \tau_1 + \cdots + p_n \leftarrow \tau_n\).

Finally, to type n-ary function application, we require the variable in function position to be an arrow type and the types of the applied terms to match the input types of the function.

We now briefly describe how we provide polymorphic types in practice, since they are heavily used in the JSkel formalization below. Polymorphism is always explicitly declared. Thus, when declaring a polymorphic type or term, one has to give its type arguments. For example, here are the declaration of the \textbf{list} type and a \textbf{map} term.
JSkel A. Khayam, L. Noizet, and A. Schmitt

<table>
<thead>
<tr>
<th>Nil</th>
<th>Cons (a, list&lt;a&gt;)</th>
</tr>
</thead>
</table>

term map<a,b>: (f: (a → b)) → (l: list<a>) → list<b> =

branch

let Nil = l in Nil<b>
or

let Cons (v, q) = l in
let w = f v in
 Cons<b> (w, q)
end

As we can see, the type arguments must be explicitly given when applying a function or a constructor. On the other hand, they are not given in a pattern, because they can always be safely inferred.

3.3 Concrete Interpretation

The concrete interpretation gives a natural, or big-step, semantics to a skeletal semantics. To define the concrete interpretation, we first explain how types are translated into sets of concrete values. Then we define the evaluation relation, which relates every term or skeleton to concrete values of their type.

Formally, for each type \( \tau \) we define the set of concrete values \( V_\tau \) in the following way.

1. If \( \tau \) is a non-specified type, \( V_\tau \) must be given.
2. If \( \tau \) is a specified type, \( V_\tau \) is the set freely generated by the constructors of type \( \tau \), where the argument of each constructor is recursively generated.
3. \( V_{\tau_1,\ldots,\tau_n} = V_{\tau_1} \times \cdots \times V_{\tau_n} \).
4. \( V_{\tau_1 \rightarrow \tau_2} = R(V_{\tau_1}, V_{\tau_2}) = \mathcal{P}(V_{\tau_1} \times V_{\tau_2}) \).

Skel being a strongly typed language, every term and every skeleton has a unique type, with the exception of empty branchings that can have any type. To recover type uniqueness, we require that empty branchings be explicitly annotated with their type. Every term of type \( \tau \) can be evaluated to zero, one, or several values of \( V_\tau \). In a sense, evaluation of a term is non-deterministic. Formally, the evaluation is a relation between terms/skeletons and values. We define in Figure 3 the evaluation of a term and of a skeleton in a given environment \( E \), which is a partial function that binds variables to values. The initial environment \( E_0 \) is an environment that binds every unspecified term to a value. We denote this evaluation as \( \Downarrow_t \) for terms and \( \Downarrow_S \) for skeletons.

To evaluate a variable, we look for its value in \( E \). To evaluate a constructor applied to a term, we evaluate the term to a value and return the constructor applied to this value. To evaluate a tuple, we evaluate each component, and return the tuple of the values. Finally, the evaluation of an abstraction is a subset of the relation \( R \) in which \( v \) relates to \( w \) if and only if the skeleton can be evaluated to \( w \), given that the abstracted variable is bound to \( v \) in the environment. It is convenient to only return a subset of the relation as we can then only consider the arguments to which the function is actually applied.

To evaluate a skeleton which is simply a term, we use the evaluation rules for terms. To evaluate a branching, we return the value \( v \) of a branch that successfully evaluates to \( v \). To evaluate the application of a variable to terms, we look up the variable in the environment and get a relation. We then evaluate the first term and check that the returned value is in the relation, which provides a result value. If the evaluation is partial and additional terms are present, this value is in turn a relation and we continue. We thus write \( R^*(v_1, \ldots, v_n) \) for
\[ E(x) = v \quad \text{VAR} \quad E, x \Downarrow_t v \quad \text{TERM} \quad E, t \Downarrow_t v \quad \text{CONST} \]

\[ E, t_1 \Downarrow_t v_1 \quad \cdots \quad E, t_n \Downarrow_t v_n \quad \text{TUPLE} \]

\[ R \subseteq \{(v, w) \mid v \in V_\tau \land (E + x \leftarrow v), S \Downarrow_S w\} \quad \text{CLOS} \quad E, \lambda x : \tau \cdot S \Downarrow_t R \quad \text{RET} \]

\[ E, (S_1 \ldots S_n) \Downarrow_S v \quad \text{BRANCH} \quad E, S \Downarrow_S v \quad (E + p \leftarrow v), S' \Downarrow_S w \quad \text{LETIN} \quad E, \text{let } p = S \text{ in } S' \Downarrow_S w \]

\[ \forall i E, t_i \Downarrow_t v_i \quad E, x \Downarrow_t R \quad R^* (v_1, \ldots, v_n, w) \quad \text{APP} \]

\[ E, (x t_1 \ldots t_n) \Downarrow_S w \]

Figure 3: Concrete Interpretation of Skeletal Semantics

\[ \exists R_1 \ldots \exists R_n, R(v_1, R_1) \land R_1 (v_2, R_2) \land \cdots \land R_n (v_n, w) \] corresponding to the curried application.

To evaluate a let binding, we evaluate the first skeleton as a value, extend the environment doing pattern matching, and evaluate the second skeleton in this extended environment. Pattern matching is recursively defined as follows.

- \( E + x \leftarrow v \) is the partial function \( E' \) s.t \( E'(x) = v \) and for all \( y \neq x \), \( E'(y) = E(y) \).
- \( E + _z \leftarrow v = E \).
- \( E + (C \ p) \leftarrow C \ v = E + p \leftarrow v \).
- \( E + (p_1, \ldots, p_n) \leftarrow (v_1, \ldots, v_n) = E + p_1 \leftarrow v_1 + \cdots + p_n \leftarrow v_n \).

To evaluate any term or skeleton, we assume given an initial environment \( E_0 \) that maps every non-specified term to a concrete value of the appropriate type.

It is easy to check that if \( \Gamma \vdash t : \tau \) and \( E, t \Downarrow_t v \in V_\tau \), then it entails that \( v \in V_\tau \), given that \( \Gamma \) matches \( E \), that is \( \forall x, \Gamma(x) = \tau \rightarrow E(x) \in V_\tau \).

4 JSkel

In this section, we present the Skel formalization of the GetValue\(^2\) method written in Skel. As a first step, we present the JSkel types and helper functions designed for a visually faithful formalization. We use a monadic approach to propagate information implicitly. We claim the result is a simple yet powerful tool.

JSkel is still work in progress. Nevertheless, we are already able to instantiate our interpreter to run basic code, namely a subset of the Expression\(^3\) grammar production, statements such as

\(^2\)https://tc39.es/ecma262/#sec-getvalue

\(^3\)https://tc39.es/ecma262/#prod-Expression

### 4.1 Towards a formalization

#### 4.1.1 ECMAScript

ECMAScript is a large vernacular specification written in an imperative style. It is divided into 28 chapters and 6 appendices. Despite its complexity and verbosity, it provides a complete specification of the behavior of JavaScript. Some choices are left to the implementation, however, so a formalization has to take into account design choices that cannot be considered standard. We explain below how we deal with them.

After an introductive part in chapters 1 to 5, where the notational and algorithmic conventions are defined, the document can be divided into three main functional groups.

Chapter 6 to 9 give a taxonomy of the ES' Data Types and Values, providing each taxon with a definition of its operations and related invariant, followed by the definition of abstract operations (type conversion, comparison, object and iterator operations), the runtime environment, and detailed classification of the type Object and its internal methods. This block of chapters gives a complete overview of the execution environment in which an ES program should be executed.

The central part of the specification, chapter 10 to chapter 16, describes the actual ES programming language. Chapters 10 and 11 focus on lexical units. Language constructs are given in chapters 12 to 15, where each construct is given with its syntactic specification and its evaluation. These describe expressions, statements, functions, and scripts. Modules are defined in chapter 16.

Chapters 17 to 27 introduce the default components of the Global Object, which can be considered as the standard library of JavaScript. In addition, a memory model is given in Chapter 28.

#### 4.1.2 Challenges of the Formalization

The first step of the mechanization in our purely functional Skel language is to deal with the imperative nature of the specification. This raises two issues.

First, there is a notion of implementation-dependent state that can be mutated. More precisely, we define by state the aggregation of all the imperative data manipulated by the specification. We design it as a record that includes the Execution Context stack, a strictness boolean flag, and a pool of Maps holding Execution Contexts, Environment Records, Realms, Script Records, and Objects. This record is left unspecified, as well as the functions to access and set it. This is representative of our general approach to have all the “implementation dependent” parts of the ES specification kept unspecified. To remain close to the imperative specification, we design a Skel state monad in Figure 4. This monad lets us implicitly pass the state around.

Second, the specification often breaks the usual control flow by having \texttt{return} in the middle of algorithms. We can capture such control flows by using nested branches (see below), but this
type state (*implementation dependent*)

\[
\text{type } st\langle\alpha\rangle := \text{state} \rightarrow (\alpha, \text{state})
\]

term st_bind\langle\alpha,\beta\rangle: (v: st\langle\alpha\rangle) \rightarrow (f: \alpha \rightarrow st\langle\beta\rangle) \rightarrow st\langle\beta\rangle = \\
\lambda s: \text{state} \rightarrow \text{let } (v', s') = v\ s \text{ in } f\ v'\ s'

term st_ret\langle\alpha\rangle: (v: \alpha) \rightarrow st\langle\alpha\rangle = \lambda s : \text{state} \rightarrow (v, s)

---

**Figure 4: State Monad in Skel**

<table>
<thead>
<tr>
<th>Field Name</th>
<th>Value</th>
</tr>
</thead>
<tbody>
<tr>
<td>[[Type]]</td>
<td>One of normal, break, continue, return, or throw</td>
</tr>
<tr>
<td>[[Value]]</td>
<td>Any ECMAScript language value or empty</td>
</tr>
<tr>
<td>[[Target]]</td>
<td>Any ECMAScript string or empty</td>
</tr>
</tbody>
</table>

\[
\text{type completionType} = \\
| Normal     | Ok \alpha |
| Break      | Abrupt \text{ion\ maybeEmpty}<\text{value}> |
| Continue   | 
| Return     | 
| Throw      | completionTarget := maybeEmpty<string> |

**Figure 5: ES Completion Record and Skel Formalization**

significantly reduces the legibility of the mechanization. We thus define a \textit{control flow} monad to simplify the mechanization.

In addition, the specification itself introduces operators that behave much like an exception monad, to deal with \textit{break}, function returns, or exceptions.

We thus propose in Section 4.1.3 an exception monad that captures the behavior of the ES’s monadic shorthands ? and !, and in Section 4.1.4 its extension to handle control-flow features.

We claim the combination of the state monad with the control-flow and the exception ones greatly simplifies our code, making JSkel easy to write, maintain and to visually compare to ES. This is shown in Section 4.2.

### 4.1.3 Completion Record and the ECMAScript Error Handling (?!?) monad

Most ES operations do not directly return values, they instead return \textit{completion records}. A Completion Record describes the runtime propagation of values and control flow. This record is composed of three fields, as depicted in Figure 5. In a nutshell, a completion record indicates what to do (this is a result, a break out of a loop, a return of a function, an exception being thrown), it contains an optional value, and in the case of a break or continue, it contains a target.

In theory, a completion record should only hold ES language values (Null, Undefined, Boolean, Number, BigInt, String, Symbol, and Object) or be empty. In practice, it is used to return many other constructions. If we consider the getValue(V) abstract operation, the completion record given as input can hold either a Value or a Reference in its [[Value]] field. Many such examples litter the spec, hence we consider completion records to be polymorphic in what their “value” field holds. We declare such completion records as \textit{type completionRecord<\alpha>}. 
Figure 6: Out and Anomaly Declarations

```ocaml
type out<α> =  
  Success completionRecord<α> 
  | Anomaly anomaly
```

Figure 7: JSkel Model of ? and !. Functions are presented in Figure B1

Figure 5 defines the types corresponding to the contents of this record: completionType holds control flow information, completionValue is either an Ok polymorphic constructor that contains the value of a non-abrupted computation, or an abruption. An ES abrupted Completion Record is one whose type is not Normal. For instance, a Throw record has an exception object as completion value, a Return record has an optional value, and a Break record has empty. Due to the aforementioned specification issues where non-values may be returned, we store the optional value in a separate constructor to be able to return it independently of the completion type. Finally, the completionTarget holds the optional string representing the target. An ES completion record for values thus has type completionRecord<maybeEmpty<value>>.

We next define a type out composed of two constructors, Success for successful computations, and Anomaly for anomalies, which intuitively corresponds to a failure of the specification. A successful computation is one that returns an ES completion record, either Normal or Abrupt. Incorrect computations are captured by the Anomaly constructor. In the specification, anomalies can be raised by assertion failures, or when it is explicitly written that an abstract operation call must not return an Abrupt. The specification authors informally guarantee that these failures never occur. We make them explicit so that we can formally express their absence and thus pave the way for a formal certification of this property.

The Anomaly constructor is of type anomaly. We define the latter with three type constructors: AbruptAnomaly that holds a completion record in case an evaluation returns an unexpected abruption, StringAnomaly that contains a textual information about an anomaly—when an assertion of the specification is broken or when an implementation-dependent operation fails, and NotImplemented that signals that we have not yet implemented some feature.

ES abruptions are propagated through an abstract method called ReturnIfAbrupt. Basi-
cally, this method gets a completion record, and either returns the value of the \([\text{Value}]\) field in case of a normal completion or propagates the abruption.

In cases of an abstract operation or a recursive evaluation, the specification uses the prefix ? to indicate that ReturnIfAbrupt has to be applied to the resulting completion. This operator is basically a monadic bind for an exception monad. The other operator used in the specification is !: it behaves like ? on a normal result, but it asserts an abruption cannot occur. We thus model it as transforming an abruption into an anomaly. These behaviors are reflected in the flow-chart presented in Figure 7, presenting respectively \(\text{out}_q\langle\alpha,\beta\rangle\) (red arrows), and \(\text{out}_b\langle\alpha,\beta\rangle\) (orange arrows). We define them as the monadic binders of the combination of \(\text{st}\) and \(\text{out}\).

\[
\text{term } (\text{out}_q|\text{out}_b)\langle\alpha,\beta\rangle : \text{st}<\text{out}<\alpha>\rightarrow (\alpha \rightarrow \text{st}<\text{out}<\beta>>) \rightarrow \text{st}<\text{out}<\beta>>
\]

\[
\text{term } \text{out}_\text{ReturnIfAbrupt}\langle\alpha,\beta\rangle : \text{out}<\alpha> \rightarrow (\alpha \rightarrow \text{st}<\text{out}<\beta>>) \rightarrow \text{st}<\text{out}<\beta>>
\]

The state monad is required for getting the result in case it is stored in the data structures in it, such as when manipulating references to values in the heap. We define, in Figure B1, getters (getNormal, getAbrupt, ...) for each completionRecord's type, and returns (retAnomaly, retAbrupt, retNormal, ...) for successful and anomaly computations. Note the use of the abruptionCoerce operation that is only defined for completion records that are abruptions. It is then the identity, but it changes the type parameter of the completion record.

Given two algorithmic steps,

1. let bar be ? AbstractOperation(foo)
2. Return bar

we represent them as

1. let\%out_q bar = abstractOperation(foo) in
2. retNormal<\tau_bar> bar

where \(\tau_{\text{bar}}\) is the type of bar.

4.1.4 A Control-Flow monad

As said earlier, the ES specification uses imperative control flow, such as returning in the middle of an algorithm. We introduce, in Figure 8, the type \(\text{controlFlow}<\alpha,\beta>\) for computations that either continue with an argument of type \(\alpha\) or that terminate with a result of type \(\beta\). It is composed of two constructors: ReturnControl, to return a result of type \(\beta\), and ContinueControl to continue the execution with a value of type \(\alpha\) to be given to the continuation.

We define a monadic binder \(\text{cf\_bind}\), a function \(\text{cf\_res}\) for extracting the value from a \(\text{controlFlow}<\alpha,\beta>\) term, and three return functions: \(\text{cf\_ret}\) for ReturnControl, \(\text{cf\_cont}\) for signaling a unit ContinueControl, and \(\text{cf\_assign}\) for an \(\alpha\) ContinueControl. We could

\[
\text{type } \text{controlFlow}<\alpha,\beta> = \begin{cases} \text{term } \text{cf\_ret}<\beta> : \beta \rightarrow \text{controlFlow}<(\cdot),\beta> \\ \text{ContinueControl } \alpha \\ \text{ReturnControl } \beta \end{cases}
\]

\[
\text{term } \text{cf\_bind}<\alpha,\beta,\gamma> : \text{controlFlow}<\alpha,\beta> \rightarrow (\alpha \rightarrow \text{controlFlow}<\gamma,\beta>) \rightarrow \text{controlFlow}<\gamma,\beta>
\]

\[
\text{term } \text{cf\_res}<\alpha,\beta,\gamma> : \text{controlFlow}<\alpha,\beta> \rightarrow (\alpha \rightarrow \beta) \rightarrow \beta
\]

Figure 8: The ControlFlow type with binders and setters
1. If foo is true Return 1
2. If bar is true Return 2
3. If baz is true Return 3
4. Return 4

(*no control-flow monad*)
branch let True = foo in 1
or let False = foo in
branch let True = bar in 2
or let False = bar in
branch let True = baz in 3
or let False = baz in 4
end

(*control-flow monad*)
let%cf_res result =
branch let True = foo in cf_ret<int> 1
or let False = foo in cf_cont<int> ()
end;%cf_bind
branch let True = bar in cf_ret<int> 2
or let False = bar in cf_cont<int> ()
end;%cf_bind
branch let True = baz in cf_ret<int> 3
or let False = baz in cf_cont<int> ()
end;%cf_bind
cf_ret<int> 4
in result

Figure 9: Code with and without Control Flow Monad. See Figure B2 for Functions

implement cf_cont using cf_assign, but we find that making the distinction explicit clarifies the code.

We illustrate in Figure 9 the translation of algorithmic steps in Skel, with and without the control-flow monad. In the examples, we use the form let (True|False) = ... in ... to test whether a value is true or false. Despite the slight overhead in notation, the control-flow approach is closer to the algorithmic steps and can be consistently applied to deal with the common case of a conditional that returns without an else branch. One can achieve a similar behavior without the control flow monad, at the cost of nested branching (highlighted on the left-hand side of Figure 9). It is possible to avoid the nesting of branching by hoisting all the branches at top-level. This results in the following code.

branch let True = foo in 1 (*1*)
or let False = foo in let True = bar in 2
or let False = foo in let False = bar in let True = baz in 3
or let False = foo in let False = bar in let False = baz in 4
end

The reason previous conditions are repeated is because there is no guarantee the evaluation of a branching will consider branches in declaration order. In addition, using a collecting semantics (where all branches are considered) would give the wrong result if we did not restate all conditions.

We define st<cf<α,out<β>>> type as the combination of the three monadic types, and accordingly, the definitions of the binders and return functions as bind, cf_out, cont, and ret. Our general approach is to encapsulate all the algorithmic steps in this type, returning an st<out<β>> at the end of every function that may raise an exception. To this end, we start each algorithm with let%cf_out result = to enter the full monad with control, and we exit the control monad at the end of the algorithm with in result.

Figure 10 gives a variant of Figure 9 with exceptions. Notice that the only thing that changes is the first step and the use of the appropriate monadic binders for type st<cf<α,out<β>>>. Despite the closeness to the algorithmic steps, the latter figure, in lines 4, 6 and 8, shows redundancy of code in the or branches. Indeed, each time the condition is not satisfied, we have to explicitly state that the evaluation continues. To avoid having to do so, we define two
1. If foo is true throw FooError
2. If bar is true Return 2
3. If baz is true Return 3
4. Return 4

let%cf_out result =
  branch let True = foo in
    let%throw fe=fooError<int> in fe
    or let False = foo in cont<int> () end;%bind
  branch let True = bar in ret<int> 2
    or let False = bar in cont<int> () end;%bind
  branch let True = bar in ret<int> 3
    or let False = bar in cont<int> () end;%bind
  ret<int> 4

in result

Figure 10: Variant of the Figure 9 with Exceptions. See Figure B3 for Functions.

boolean binder terms, ifpTrue and ifpFalse, for introducing partiality in branchings. The
ifpTrue binder, in the case v is True, applies the bind function f to a unit, otherwise, in the
case v is False, the branch returns a continuation of type α. The ifpTrue term definition, and
its application to the Figure 10 example, results in the following code.

```
term ifpTrue<α>:
  (v: boolean) →
    (f: () → st<controlFlow<(),out<α>>>
      ) →
    st<controlFlow<(),out<α>>> =
    branch let True = v in f ()
    or let False = v in cont<α>()
    end
```

```
let%cf_out result =
  branch foo;%ifpTrue
    let%throw fe=fooError<int> in fe end;%bind
  branch bar;%ifpTrue ret<int> 2 end;%bind
  branch bar;%ifpTrue ret<int> 3 end;%bind
  ret<int> 4

in result
```

4.2 The GetValue(V) example

In this section, we illustrate our design choices through the mechanization in Skel of the
GetValue abstract method. This method, presented in Figure 11, is one of the most referenced
methods in the specification, as it is often called after the evaluation of syntactic constructors of
the language. It takes V as input, a completion record containing either a value or a reference,
and it returns a value as output. In a nutshell, if V is a value, GetValue returns it, and if
it is a reference, GetValue acts like a binding resolver to obtain a primitive value, Object, or
Environment Record. The reference type is shown in Figure 12. It is a record with a field
[[BaseValue]] of type environmentRecord or value, and additional fields not relevant here.
Our mechanization in Skel leaves this type unspecified. We describe the types and terms used
in our formalization of getValue in Figure B4.

We now describe step by step how we formalize this method (the code is collected as a
single function in Figure B5). The type mixing values and references is the specified type
valref defined as Value value | Reference reference. The argument of GetValue thus has type out<valref>
. The first step of the abstract method applies ReturnIfAbrupt on V. In case it is an abruption, the method propagates the completion record to the caller, otherwise it
extracts the completionValue. To model this, we use the %returnIfAbrupt monadic binder.

```
1. let%returnIfAbrupt v = v

If V is a normal completion, then the newly bound V will have type valref. We could use a
different type for returnIfAbrupt to be able to write let%bind v = returnIfAbrupt v, but
we would then need to specify the polymorphic type components of returnIfAbrupt. We are
working on type inference to enable such a change in the future.
1. ReturnIfAbrupt(V).
2. If Type(V) is not Reference, return V.
3. Let base be GetBase(V).
4. If IsUnresolvableReference(V) is true, throw a ReferenceError exception.
5. If IsPropertyReference(V) is true, then
   a. If HasPrimitiveBase(V) is true, then
      i. Assert: In this case, base will never be undefined or null.
      ii. Set base to ToObject(base).
   b. Return ? base.[[Get]](GetReferencedName(V), GetThisValue(V)).
6. Else,
   a. Assert: base is an Environment Record.
   b. Return ? base.GetBindingValue(GetReferencedName(V), IsStrictReference(V)).

Figure 11: The ECMAScript’s GetValue(V)

<table>
<thead>
<tr>
<th>Field Name</th>
<th>Value</th>
</tr>
</thead>
<tbody>
<tr>
<td>[[BaseValue]]</td>
<td>Value or Environment Record</td>
</tr>
<tr>
<td>[[Strict]]</td>
<td>Boolean Flag</td>
</tr>
<tr>
<td>[[ThisValue]]</td>
<td>Value or Empty</td>
</tr>
<tr>
<td>[[Name]]</td>
<td>String Value</td>
</tr>
</tbody>
</table>

Figure 12: The Reference type

The next step inspects the type of V. If it is not a reference, hence it is a value, we return it. Otherwise, the ifpFalse implicitly propagate a continuation.

2 branch valref_Type(v, T_Ref);%ifpFalse let Value v = v in ret<value> v end;%bind

The GetBase method in the third instruction takes the [[BaseValue]] from the reference V, and assigns it to base. The previous step guarantees that V has type reference, so we procede:

3 let Reference v = v in let base = getBase(v) in

Now base has type ref bv. The step 4 follows the same pattern as step 2:

4 branch isUnresolvableReference v;%ifpTrue
5 let%throw refErr = referenceError<value> in refErr end;%bind

An unresolvable reference is the one that has Undefined as [[BaseValue]]. In this case, a ReferenceError is thrown. Otherwise, step 5 inspects whether V is a property reference. Being a property reference means that a reference have a non-null, defined base value of type value.

6 branch let True = isPropertyReference v in

Step 5.a. checks whether the reference V has a primitive [[BaseValue]]. Having a primitive base means that the reference points to a primitive value, i.e., a value of type Boolean, String, Symbol, BigInt, or Number. After some assertions, this algorithmic step sets base as ToObject
applied to base. Note the use of the \%b monadic operator corresponding to the ! in the call: ToObject must return a normal result, otherwise an anomaly is raised. There is no else branch, but since base is not primitive in that case, it has to be an Object. We extract it using pattern matching. In both case, the resulting loc_Object is returned to the binder let%bind base using assign.

```
let%bind base =
branch let True = hasPrimitiveBase v in let R_Value base = base in
val_Type(base,T_Null);%assF
let%b base = toObject(base) in assign<loc_Object,value>(base)
or let False = hasPrimitiveBase v in
let R_Value (Obj base) = base in assign<loc_Object,value>(base)
end in
```

In line 9, we use the monadic binder assF\langle value \rangle for “assert false”. If the result of the val_type application is true, an anomaly is raised.

Once base is set, the algorithmic step 5.b calls the object internal method [[Get]]. The JSkel formalization is straightforward with it, using base as the first argument to o_Get.

```
let name = getReferencedName(v) in let%q thisVal = getThisValue(v) in
let%q v' = o_Get(base, name, thisVal) in ret<value> v'
```

Note that we make explicit the order of the calls to getReferencedName and getThisValue, as Skel requires function application to be to fully computed terms. We could faithfully model different execution orders using branching, but this would make the mechanization confusing. The call to getReferencedName is pure, although the specification contains an assertion\(^8\). Since we know that the assertion is satisfied by typing, we do not include it. The call to getThisValue, however, is not pure as the assertion there\(^9\) is not captured by typing (although we have checked that the reference is indeed a property reference at step 5). We thus use the %q binder in that case to extract the pure value or propagate the anomaly, if any.

If V is not a property reference, then base must be an Environment Record. In this last step, we check that this is the case using an assertion, then we apply the getBindingValue operation.

```
or let False = isPropertyReference v in ref_Type(base, T_R_EnvRec);%assT
let R_EnvironmentRecord base = base in
let name = getReferencedName(v) in let strict = isStrictReference(v) in
let%q v' = er_GetBindingValue(base, name, strict) in ret<value> v'
```

As with most of the algorithms in our mechanization, we need to return a result that is out of the control monad. The whole algorithm is thus surrounded by the following piece of code.

```
let%cf_out result = (*GetValue’s algorithmic steps*) in result
```

5 Conclusion and Future Work

We have described how the Skel language can be used to mechanize complex semantics. In particular, we have shown how carefully chosen monads can help in keeping the specification and

---

\(^8\)https://tc39.es/ecma262/#sec-getreferencedname
\(^9\)https://tc39.es/ecma262/#sec-getthisvalue
its mechanization very similar. The current state of our formalization of JavaScript can be found online, at https://gitlab.inria.fr/skeletons/jskel. It includes the Skel formalization of the language, boilerplate code to integrate with existing JavaScript parsers (see Appendix A), and an implementation of all the unspecified terms. Using necro-ml [7], we have all that is needed to generate an executable semantics.

To this point, our main effort has been in mechanizing the foundations of the language, of which GetValue is a good example. We are now ready to formalize more constructs than simple expressions and statements.

We also have been able to start the evaluation of maintainability of the approach. This project began with the formalization of ECMAScript 2020, but after few months, we switched to the newer ECMAScript 2021. This process took only a day of work. This may change once we have a more complete formalization of the specification, but as we can easily produce a list of differences between specifications, the amount of work is proportional to the size of changes and not to the size of the specification. In particular, we only have one mechanization to change, instead of two formal developments and a correctness proof as is the case for JSCert.

A short term goal is to formalize enough of JavaScript to have all the features needed to run the ECMA-262 test suite. A longer term goal is to validate that we can use our mechanized semantics to prove properties of the language. A first property of interest is the guarantee that assertions in the specification are actually satisfied. To achieve this goal, we will use the necro-coq tool [8]. As the formalization of JavaScript helped us refine the Skel language, we believe the generation of a Coq semantics will provide many opportunities to improve necro-coq.

References

A The parsing process

In JSkel, we formalize the JavaScript syntax as it is defined in ECMAScript. For instance, we give in Figure A1 the description of the Expression type, which is the same as its ES counterpart.

Unfortunately, no existing parser of JavaScript provides such a faithful representation. We thus had to choose between implementing our own parser or translating the AST provided by on-the-shelf parsers. We chose the latter.

More precisely, we use a parser that conforms to the SpiderMonkey’s Parser API, which is followed by all parsers we have found. We chose the Flow Parser library because it is written in OCaml, which is the language we can instantiate our interpreter into, and because it is used to manipulate JavaScript in industrial setting.

When instantiating the interpreter, we define a transformation that, given a Flow AST, produces a well-typed ES AST. In the long term, we would like to write a parser faithful to the specification, but the complexity of ES syntax makes it challenging.

In Figure A2, we show the resulting AST of a program produced by the Flow Parser. Its transformation is presented in Figure A3.

An execution of our interpreter on a JS program first performs InitializeHostDefinedRealm, then parses the source code, and finally calls ScriptEvaluation.

---

Figure A1: Type definition of expressions in JSkel

type expression =
| Expr_AssExpr assignmentExpression
| Expr_Comma (expression, assignmentExpression)

Figure A2: Parsing of “let x = 1; x + 1”
B Unspecified Terms

For introductive purposes to JSkel, we have left these terms unspecified, presenting only their definition. All the terms presented in this section are fully implemented in the skel language.

```plaintext
(*Operations of the completionRecord type*)

term getNormal<α> : completionRecord<α> → α

term abruptCoerce<α, β> : completionRecord<α> → completionRecord<β>

(*out type getters and setters*)

term getAnomaly<α> : out<α> → anomaly

term retAnomaly<α> : anomaly → st<out<α>>

term retAbrupt<α> : (maybeEmpty<value>, completionType) → st<out<α>>

term retNormal<α> : α → st<out<α>>
```

Figure B1

Figure A3: The transformation of Figure A2
(*Types*)

type boolean =
| False
| True

(*Binders*)
term cf_bind<α, β, γ> :
controlFlow<α, β> →
(α → controlFlow<γ, β>) →
controlFlow<γ, β>
term cf_res<α, β> :
controlFlow<α, β> →
(α → β) →
β

(*Returns*)
term cf_ret<α> : α → controlFlow<st<out<α>>>
term cf_cont<α> : () → controlFlow<st<out<α>>>

Figure B2

(*Binders*)
term cf_out<β, γ> :
st<controlFlow<()>, β> →
(β → γ) →
st<γ>
term throw<β> :
(() → st<out<β>>>) →
(st<controlFlow<()>, out<β>>> → st<controlFlow<()>, out<β>>>) →
st<controlFlow<()>, out<β>>>
term bind<α, β, γ> :
st<controlFlow<α, out<β>>> →
(α → st<controlFlow<γ, out<β>>>) →
st<controlFlow<γ, out<β>>>

(*Returns*)
term cont<β> : () → st<controlFlow<()>, out<β>>>
term ret<β> : β → st<controlFlow<()>, out<β>>>

(*Functions*)
term fooError<α> : () → st<out<α>>

Figure B3
(*Types*)

```plaintext
type valref = | Reference reference | Value value

type type_valref = | T_Ref | T_Val

type ref_bv = | R_Value value | R_EnvironmentRecord loc_EnvironmentRecord

type type_ref_bv = | T_R_EnvRec | T_R_Value

type value = | Null | Undefined | String string | Number number | BigInt bigInt | Object loc_Object | Symbol value | Boolean boolean

type type_value = | T_Null | T_Undefined | T_String | T_Number | T_BigInt | T_Object | T_Symbol | T_Boolean
```

(*Binders*)

```plaintext
term returnIfAbrupt<α, β, γ> : out<α> → (α → st<controlFlow<γ, out<β>>>) → st<controlFlow<γ, out<β>>>

term assF<α, β> : boolean → (boolean → st<controlFlow<α, out<β>>>) → st<controlFlow<α, out<β>>>

term assT<α, β> : boolean → (boolean → st<controlFlow<α, out<β>>>) → st<controlFlow<α, out<β>>>

(*Returns*)

```plaintext
term assign<α, β> : α → st<controlFlow<α, out<β>>>
```

(*Boolean type checkers*)

```plaintext
term valref_Type : (valref, type_valref) → boolean

term val_Type : (value, type_value) → boolean

term ref_bv_Type : (ref_bv, type_ref_bv) → boolean
```

(*Reference type abstract methods*)

```plaintext
term hasPrimitiveBase : reference → boolean

term getBase : reference → ref_bv

term isUnresolvableReference : reference → boolean

term isPropertyReference : reference → boolean

term getReferencedName : reference → string

term isStrictReference : reference → boolean
```

(*Type conversion operations*)

```plaintext
term toObject : value → loc_Object
```

(*Object internal method*)

```plaintext
term o_Get : (loc_Object, string, value) → st<out<value>>
```

(*Environment Record’s abstract method*)

```plaintext
term er_GetBindingObject : (loc_EnvironmentRecord, string, boolean) → st<out<value>>
```

(*Errors*)

```plaintext
term referenceError<α> : () → st<out<α>>
```

Figure B4
term getValue : (v : out<valref>) -> st<out<value>> =

    let%cf_out result =
    let%returnIfAbrupt v = v in
    branch valref_Type(v, T_Ref);%ifpFalse let Value v = v in ret<value> v end;%bind
    let Reference v = v in let base = getBase(v) in
    branch isUnresolvableReference v;%ifpTrue
      let%throw re = referenceError<value> in re end;%bind
    branch let True = isPropertyReference v in
      let%bind base =
        branch let True = hasPrimitiveBase v in let R_Value base = base in
        val_Type(base, T_Null);%assF val_Type(base, T_Undefined);%assF
        let%b base = toObject(base) in assign<loc_Object,value>(base)
or let False = hasPrimitiveBase v in
        let R_Value (Obj base) = base in assign<loc_Object,value>(base)
    end in
    let name = getReferencedName(v) in let%q thisVal = getThisValue(v) in
    let%q v' = o_Get(base,name,thisVal) in ret<value> v'
or let False = isPropertyReference v in ref bv_Type(base,T_R_EnvRec);%assT
    let R_EnvironmentRecord base = base in
    let name = getReferencedName(v) in let strict = isStrictReference(v) in
    let%q v' = er_GetBindingValue(base,name,strict) in ret<value> v'
end

in
result

Figure B5