Gradual Typing Embedded Securely in JavaScript

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Abstract

JavaScript's flexible semantics makes writing correct code hard and writing secure code extremely difficult. To address the former problem, various forms of gradual typing have been proposed, such as Closure and TypeScript. However, supporting all common programming idioms is not easy; for example, TypeScript deliberately gives up type soundness for programming convenience. In this paper, we propose a gradual type system and implementation techniques that provide important safety and security guarantees.

We present TS*, a gradual type system and source-to-source compiler for JavaScript. In contrast to prior gradual type systems, TS* features full runtime reflection over *three* kinds of types: (1) simple types for higher-order functions, recursive datatypes and dictionary-based extensible records; (2) the type `any`, for dynamically type-safe TS* expressions; and (3) the type `un`, for untrusted, potentially malicious JavaScript contexts in which TS* is embedded. After type-checking, the compiler instruments the program with various checks to ensure the type safety of TS* despite its interactions with arbitrary JavaScript contexts, which are free to use `eval`, stack walks, prototype customizations, and other offensive features. The proof of our main theorem employs a form of type-preserving compilation, wherein we prove all the runtime invariants of the translation of TS* to JavaScript by showing that translated programs are well-typed in JS*, a previously proposed dependently typed language for proving functional correctness of JavaScript programs.

We describe a prototype compiler, a secure runtime, and sample applications for TS*. Our examples illustrate how web security patterns that developers currently program in JavaScript (with much difficulty and still with dubious results) can instead be programmed naturally in TS*, retaining a flavor of idiomatic JavaScript, while providing strong safety guarantees by virtue of typing.

Categories and Subject Descriptors D.3.1 [*Programming Languages*]: Formal Definitions and Theory—Semantics; D.4.6 [*Operating Systems*]: Security and Protection—Verification.

Keywords type systems; language-based security; compilers

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1. Introduction

Writing secure JavaScript is notoriously difficult. Even simple functions, which appear safe on the surface, can be easily broken. As an illustration, consider the script below, simplified from actual scripts in the OWASP CSRFGuard (2010) and Facebook API (2013), security-critical JavaScript libraries that aim to impose access controls on the networking APIs provided by web browsers.

```
function send(u,m) {/*Stand-in for XMLHttpRequest*/}
function protect(send) {
  var whitelist={"https://microsoft.com/mail":true,
                 "https://microsoft.com/owa":true};
  return function (url, msg) {
          if (whitelist[url]) send(url, msg);
        };
} send = protect(send);
```

The function call `protect(send)` on the last line returns a function that interposes an access control check on `send`. If this script were to run in isolation, it would achieve its intended functionality. However, JavaScript programs rarely run in isolation—programmers explicitly link their code with third-party frameworks, and, worse, unexpected code fragments can be injected into the web sandbox by cross-site scripting attacks. For example, the following script running in the same sandbox as `protect` could succeed in sending a message to an unintended recipient.

```
Object.prototype["http://evil.com"]=true;
send("http://evil.com", "bypass!");
```

This is just one attack on `protect`; similar attacks are often the consequence of unintended callbacks to untrusted code (caused, for example, by implicit coercions, getters or setters, prototype hacking, and global object overwriting). Experimentally, we found and reported several such security flaws in the OWASP CSRFGuard and Facebook API scripts, suggesting that untrusted callbacks remain difficult both to prevent and to contain.

1.1 Attacks $\approx$ Type Errors Arguably, each of these attacks can be blamed on JavaScript's lax notion of dynamic type-safety. Many runtime behaviors that are traditionally viewed as dynamic type errors (e.g., accessing non-existent properties from an object) are not errors in JavaScript. However, almost *any* such error while running a sensitive script can be exploited by a malicious context, as follows. Anticipating that the script will dereference an otherwise undefined property `x`, which in JavaScript would just return the value `undefined`, a hostile context can define

```
Object.defineProperty(Object.prototype,"x",
                      {get:function(){/*exploit*/}});
```

then run the script and, as `x` is dereferenced and triggers the callback, access any argument on the caller stack. Thus, for protecting good scripts from bad ones, silent type errors in JavaScript are just as dangerous as buffer overruns in C. Despite the numerous dynamic checks performed by the JavaScript runtime, some stronger notion of type safety is called for.

Other researchers have made similar observations. For example, Fournet et al. (2013) show several attacks on a similar piece of code and argue that carefully mediating the interaction between a script and its context is essential for security. They show how to compile f*, a statically typed, ML-like source language to JavaScript in a fully abstract way, allowing programmers to write and reason about functions like `protect` in ML, and to let their compiler generate secure JavaScript automatically. This is an attractive design, but the sad truth is that millions of JavaScript programmers are unlikely to switch to ML. Meanwhile, Bhargavan et al. (2013) have developed DJS, a minimal, statically typed, secure core of JavaScript, primarily for writing first-order, string-processing functions using arrays of fixed size. DJS is suitable for writing security-critical code, like cryptographic libraries, and Bhargavan et al. prove that the behavior of programs accepted by the DJS type checker is independent of the JavaScript environment in which they may run. Many others propose analyses for subsets of JavaScript (Chugh et al. 2012; Guarnieri and Livshits 2009; Guha et al. 2011; Hedin and Sabelfeld 2012; Swamy et al. 2013; Taly et al. 2011), although the guarantees provided only hold if the entire environment in which the program runs is also in the subset in question. Efforts like AdSafe (`http://www.adsafe.org`) and Caja (`http://code.google.com/p/google-caja`) address the problem of safely loading untrusted JavaScript into a program's environment after analyzing and rewriting it. However, in order to do this reliably, various JavaScript features are forbidden, e.g., Politz et al. (2011) show that by ruling out many commonly used JavaScript features (e.g., `eval` or explicit prototype manipulation), AdSafe can safely sandbox untrusted code. Although we find the work on JavaScript subsets a promising line to pursue (indeed, even the ECMAScript standard is evolving towards the definition of subsets, such as the strict mode), assumptions about subsetting the environment are hard to validate in the presence of cross-site scripts.

1.2 TS*: a gradually type-safe language within JavaScript

This paper presents TS*, a source programming language that retains many of the dynamic programming idioms of JavaScript, while ensuring type-safety even in an untrusted JavaScript environment. TS* supports writing functions like `protect` exactly as shown, while a compiler from TS* to JavaScript ensures that the access control check in `protect` cannot be subverted.

Although significantly more flexible than f* or DJS, TS* still rules out many inherently unsafe features of JavaScript for the code we protect, thereby enforcing a stricter programming discipline and facilitating security code reviews. We intend TS* to be used to protect security-critical scripts—guarding, for instance, sensitive resources and capabilities—executed on web pages that also include dynamically loaded, untrusted, potentially malicious scripts. By necessity, most of the code running on these pages is provided by third parties; we leave such code unchanged (and unprotected). Nonetheless, by lightly rewriting the security-critical scripts, and gradually typing them (Siek and Taha 2006), we enforce a strong notion of dynamic type-safety. In addition, we protect our scripts using type-directed wrappers to shield them from adversarial JavaScript environments. This places them on a robust type-safe foundation, from which we can reason about their security. In comparison, other gradually typed extensions to JavaScript, like TypeScript (`http://www.typescriptlang.org`) and Closure (`https://developers.google.com/closure/compiler`), aim primarily to increase programmer productivity by using types for documentation and code completion, but provide no safety guarantee.

The concrete syntax of TS* extends JavaScript with a type-annotation language based on TypeScript. Our implementation works by first using TypeScript to infer types for every sub-term. It then uses the inferred types to type-check the program once again, this time using a more restrictive type system. If this step is successful, the program is compiled to plain JavaScript (instrumented with various runtime checks to enforce type-safety) for execution. The TS* type system itself has the following prominent features:

A statically typed core of functions, datatypes and records The base type system of TS* includes primitive types like `bool`, `number` and `string`; higher-order function types; recursive datatypes; and extensible records of fields with optional immutability annotations. Records are equipped with a structural subtyping relation. For example, the type `point` defined below is a subtype of a record that omits some of its fields, and function subtyping is, as usual, contravariant on the arguments and covariant on the results.

```
{x:number;  y:number;  setX:const (number->unit)}
```

Dynamically typed fragment The type `any` is for dynamically typed TS* expressions. All the types from the statically typed core are subtypes of `any`, and all TS* terms whose subterms all have type `any` can always be given the type `any`. In the spirit of JavaScript, in the `any`-fragment, we view records as extensible dictionaries with string-valued keys. TS* supports the use of computed properties, e.g., in `any`-typed code, expressions like `p["set"+"X"]` or `whitelist[url]` are legal ways to safely project the appropriate field from the underlying object, if it exists. Dynamically typed code and its interactions with statically typed code are instrumented by our compiler for safety. As far as we are aware, ours is the first gradual type system to soundly support dictionary-based mutable records and computed properties.

Runtime reflection over types Case analysis on the runtime type of a value is a common idiom in JavaScript and other dynamically typed languages—Guha et al. (2011) present several typical uses of JavaScript's `typeof` operator. TS* embraces this idiom and compiles programs with runtime type information (RTTI) to support introspection on *all source types* at runtime, e.g., `isTag<point>p` checks whether the RTTI of `p` is a structural subtype of `point`. In addition to providing an expressive source programming construct, RTTI also forms the basis of an efficient and simple enforcement mechanism for gradual typing, an alternative to prior proposals based on higher-order contracts (Findler and Felleisen 2002).

un*, the type of the adversary, mediated by wrappers* Finally, and most distinctively, TS* provides a *second dynamic type*, un. Any JavaScript term which cannot be typed either statically or dynamically can be given the type un and simply passes through our compiler without further analysis or modification. As such, un is the type of arbitrary, potentially adversarial JavaScript expressions. Our un type is reminiscent of types for adversaries, as proposed by Gordon and Jeffrey (2001). However, unlike prior uses of un in the context of secure compilers (e.g. Fournet et al. 2013), un is a first-class type in TS*: un values may be stored in records, used as arguments and results of functions, etc. The type un is incomparable to any in the subtyping relation and, in contrast with any, all operations on un values are mediated by wrappers that safely build coercions to and from un (as well as other types). The wrappers enforce a strict heap separation between the un-context and typed TS*, ensuring that adversarial code cannot break type invariants.

1.3 Evaluating TS*: theory and practice We specify our compiler as a type-directed translation relation (§3). To formalize properties of the translated program, we give TS* a translation semantics to JS*, a dependently typed model of JavaScript developed by Swamy et al. (2013), which is in turn based on λJS by Guha et al. (2010). Precise monadic refinement types in JS* allow us to conveniently phrase our metatheory (§4) in terms of type-correctness of JS*, yielding three main properties:

Memory isolation: un-typed code (i.e., the adversary) cannot directly call, read, write, or tamper with typed values.

Static safety: statically typed code is safely compiled without any runtime checks, even in the presence of type-changing updates.

Dynamic safety: runtime type information is sound and at least as precise as the static type.

Experimentally, we evaluate TS* by programming and adapting several security-sensitive JavaScript web libraries (§6). Our examples include a OWASP reference library to protect against cross-site request forgeries (CSRF) (Barth et al. 2008); and an adaptation of secure login and JSON-validation scripts within the Facebook API. In summary, our main contributions include:

(1) a type system and runtime support for safely composing statically typed, dynamically typed, and arbitrary JavaScript; (§3)

(2) a type safety theorem and its proof by translation to JS*; (§4)

(3) a prototype implementation, including a protocol to ensure that our runtime support runs first on pages hosting compiled TS*, and securely initializes our type invariant; (§5)

(4) security applications, illustrating a series of authorization and access control patterns taken from popular security-sensitive web applications and libraries, motivated by new attacks. (§6)

As such, TS* is the first system to provide a notion of type safety useful for secure programming, while handling *all* of JavaScript.

The latest version of our compiler, programming examples, attacks, sample web-application deployments, and a technical report with the full formalization and proofs are all available from `http://research.microsoft.com/fstar`.

2. An overview of TS*

We begin by presenting the design of TS* informally, using several small examples for illustration. We use the concrete syntax of TypeScript with minor simplifications; for instance we write function types as `t -> t'` rather than TypeScript's `(x:t) => t'`. We also extend the syntax of TypeScript with datatype declarations and immutability qualifiers.

2.1 Gradually securing programs by moving from un to any

While we envisage TS* as the basis of a full-fledged gradually typed web-programming language, we initially consider JavaScript programmers willing to harden safety- and security-critical fragments of their code. They can start by giving their existing JavaScript code the type un in TS*, and then gradually migrating selected fragments to (dynamically) type-safe code in TS* using the type any.

This exercise is valuable since any code in TS* enjoys a *memory isolation* property, a robust foundation upon which to build secure sub-systems of a larger program. Memory isolation alone prevents many common attacks. For example, the prototype poisoning attack of §1 occurs because of a failure of memory isolation: the command `whitelist[url]` causes a prototype-chain traversal that ends with reading a field of `Object.prototype` which, unfortunately, is a reference to an object controlled by the adversary. Passing `protect` (unchanged) through the TS* compiler while giving `whitelist` the type `any` produces an instrumented version of `protect` with checks to prevent any property dereferences from `whitelist` from accessing `Object.prototype` (or any other un object), thereby foiling the attack. Specifically, from memory isolation, we can prove that every dereference of a field of any object in TS* will only read the immediate fields of that object, and will never access a prototype controlled by the adversary. This ensures that `whitelist[url]` returns true only if `url` is immediately defined in `whitelist`.

Undecorated TS* programs can generally be given the type any (as long as they are well-scoped). Every function parameter in an unannotated TS* program defaults to the type any; every `var`-bound variable is given the type of its initializer. Under this convention, in the program from §1, the type of `protect` is `any -> (any,any)-> any`, which is a subtype of any. When deploying a TS* program, we assume that the JavaScript global object (the `window` object in most browsers) and all objects reachable from it are under control of the attacker. Thus, it is not safe to simply store `protect(send)` into `window.send`, since that would break memory isolation and leak a value of type `any` to un-safe code—our type system prevents the programmer from doing this by mistake.

Instead, TS* provides wrappers to safely export values to the context. The TS* expression `wrap<un>(protect(send))` wraps the `protect(send)` closure and yields a value of type un, indicating that it is safe to hand to any JavaScript context while preserving memory isolation. Dually, for `e:un`, the expression `wrap<any>(e)` safely imports `e` from the context and gives it the type any.

Providing a JavaScript implementation of `wrap` is non-trivial. We base our implementation on wrappers defined by Fournet et al. (2013). Their wrappers are designed to safely export statically typed values from the translation of an f* program (roughly, a simply typed subset of ML) to its JavaScript context; and to import untyped values from the context into f* at specific types. For example, Fournet et al.'s $\texttt{down}_{(t*u)}$ exports a pair of translated f* values (v_1,v_2) of type $(t*u)$ to the context, by building a new object with two fields initialized to $\texttt{down}_t(v_1)$ and $\texttt{down}_u(v_2)$. A corresponding wrapper $\texttt{up}_{(t*u)}$ does the converse, safely copying a pair from the context and building a value that is the translation of an f* pair of type $(t*u)$. Fournet et al. provide $\texttt{up}_t$ and $\texttt{down}_t$ wrappers for types t including unit, bool, string, number, pairs, recursive datatypes, and functions. We extend their constructions to additionally build wrappers to and from the type any, and to handle cyclic data structures that can be constructed using records with mutable fields.

To illustrate wrappers in action, we elaborate on our first example. Suppose we wished to protect `window.send` (a fictitious but simpler stand-in for JavaScript's `XMLHttpRequest` object) with an access control check. To support this, the standard library of TS* provides a facility to read fields from, and write fields to, the global object by including the following safe interface to the `window` object implemented in JavaScript. The `win` object, whose interface is shown partially below, shadows the fields of the `window` object, safely reading and writing it within a wrapper to ensure that the `attacker`-controlled window does not break type safety. Using `win` within a TS* program, we can safely import `window.send`, protect it, and export it back to the context using the following snippet of code, typed in a context where the `win` object has mutable un-typed fields. Of course, the attacker may *a priori* obtain a copy, and even redefine `window.send` before our code has the chance to protect and update it, but this is an orthogonal problem, solved once for all TS*programs—§2.3 and §5 present our mechanisms to ensure that our scripts run first on a web page.

```
interface win {send: un; ... }
win.send = wrap<un>(protect(wrap<any>(win.send)));
```

Wrappers are expensive, since they deeply copy the contents of objects back and forth, and—by design—they are not necessarily semantics-preserving. (For instance, they sanitize values, filter out some properties, and prevent some aliasing.) Thus, in an attempt to minimize the amount of copying, we may rewrite `protect`, by adding a few types, as shown below.

```
function protect(send:(string,un) -> un) {
  var whitelist={"http://microsoft.com/mail":true,
                 "http://microsoft.com/owa":true};
  return function (url:string, msg:un) {
            if (whitelist[url]) send(url, msg); };
};
win.send =
  wrap<un>(protect(wrap<(string,un)->un>(win.send)));
```

Intuitively, the `msg` argument in the closure returned by `protect` is treated abstractly, that is, our script does not directly access it. Thus, there is no need to import that argument from the context (potentially performing a deep copy). On the other hand, the `url` argument is not abstract—it is used to project a field from the whitelist, and, as such, it had better be a string. The type system of TS* gives us the flexibility to express exactly what should be imported from the context, helping us find a good balance between security and performance. The explicit use of un and `wrap` are advances of TS* relative to prior languages such as f* or, for that matter, any prior gradually typed programming language.

2.2 Expressing invariants with assertions over runtime types

As can be expected of gradual typing, a TS* program migrated from un to any can then, with some effort, be made increasingly statically typed. Static types can improve runtime safety, robustness of code, modularity, as well as provide better IDE support. Static types in TS* also improve performance relative to any-typed code and, relying on RTTI, can enforce data invariants. This is enabled by *static safety* and *dynamic safety*, two properties (in addition to *memory isolation*) provided by TS*.

Static safety TS* ensures that, at runtime, no failures happen during the execution of statically typed parts of the source program. Since there are no runtime checks in the compiled JavaScript for such parts, as a bonus, the performance of statically typed TS* code will approach that of native JavaScript (and potentially exceed it, if the type information can be communicated to the VM).

Dynamic safety Every TS* value $v : t$ (where $t \neq$ un) is compiled to JavaScript with runtime type information (RTTI) that initially reflects v's static type t. TS* ensures that, while v's RTTI may evolve during execution (e.g., as fields are added to an extensible record), it is always (a) a subtype of v's static type t, and (b) a sound approximation (supertype) of v's most-precise current type, i.e., the RTTI of v may evolve only monotonically with respect to the subtyping relation. We call this property *dynamic safety*.

As an illustration, consider the example below, which codes up a lightweight form of objects with extensible records and closures in TS*, where `point` is the type defined in §1.

```
function Point(x, y) {
  var self = {};
  self.x=x;
  self.y=y;
  self.setX=function(d:number) { self.x = d; };
  return setTag<point>(self);
}
```

The function `Point` creates a new point. It allocates a new empty record and stores it in the local variable `self`, then it adds three fields `x`, `y`, and `setX`. The static type of `self` is just the empty record. However, TS* allows us to add more fields to `self` than those documented in its static type. As such, the static type of a record only describes a subset of the fields in the term, as is usual with width-subtyping. Deleting fields from records is also possible—we discuss this in more detail in §3.

In the last line of `Point`, `setTag<point>(self)` checks at runtime if the content of `self` is compatible with the `point` type, and fails otherwise. The term `setTag<point>(self)` has static type `point`, although the static type of `self` remains unchanged.

Assertions like `setTag` allow source programmers to safely update RTTI while maintaining the runtime type invariant. Once a value has been tagged as a `point`, then it is guaranteed to always remain a point. A programmer may choose to add fields to a `point`, or to further update its type information (e.g., turning it into a `coloredPoint`), but it will always contain at least the fields of a `point`. Any attempt to delete, say, the `x` field, or to change it in a type-incompatible way (using for instance a dynamically typed alias to the point) will cause a runtime error.

In contrast, statically typed code raises no such errors. TS* infers that function `Point` has type `(any,any)-> point`, so the code below is statically type safe and does not require any runtime checks.

```
var o = Point(0,0); o.setX(17);
```

As another example,[1] consider that popular web frameworks, like Dojo, provide implementations of the JSON Schema standard.[2] This allows programmers to validate JSON data, writing verbose schemas for them also as JSON objects. In TS*, data invariants can be expressed and enforced directly using types, rather than via schemas. For example, to check that a JSON string can be parsed into an array of user identities, we can write the TS* code below, assuming that `JSON.parse` has type `string -> any`. (See online materials for a TS* implementation of a JSON parser.)

```
type users = {id:number; user:string}[]
function check(j:string) : users {
  var o = JSON.parse(j);
  if (canTag<users>(o)) return setTag<users>(o);
  else return [];
}
```

The schema is captured by the type `users`. We parse a string `j` as JSON using `JSON.parse`, then use the TS* operator `canTag<t>(o)` to check that `o`'s contents are consistent with `t`. If the check succeeds, we stamp `o` as a valid `users` object and return it.

2.3 Reliable primitive operations

Since the global `window` object is shared with the adversary, all objects reachable from `window` may be compromised. This includes all built-in objects provided by the VM, e.g., `Object.prototype`, `Array.prototype`, the default `String` object, and others. In order to ensure memory isolation, translated TS* programs should never read from any of those objects. This is remarkably difficult to arrange in JavaScript, since several primitive operations, e.g., reading and writing fields, depend on base prototypes, as illustrated in §1. Thus, in the face of attacker-controlled prototypes, even simple manipulations of objects are unreliable.

Thankfully, there is a relatively simple way out. JavaScript (since ECMAScript 5) provides a function `Object.create`, such that `Object.create(arg)` creates a new object with its prototype field initialized to `arg`. As a special case, `Object.create(null)` creates an object with a `null` prototype, meaning that the created object *does not* inherit the attacker-controlled `Object.prototype`. Our compiler uses `Object.create(null)` to create new objects in the translation of a TS* program, allowing manipulations of object fields simply as `o.f=v` or `o["f"]=v`, without worrying that a traversal of `o`'s prototype chain will reach an attacker-controlled object.

There is one exception, however. Function objects cannot be allocated using `Object.create`—they must inherit from the attacker-controlled `Function.prototype`. However, by an invariant of the translation, we can ensure that the only field ever accessed of a non-un-typed function `f` is "`rtti`". So, as long as `Function.prototype` contains a safe, immutable `rtti` field, accessing `f.rtti` will never trigger adversarial code.

The correctness of our translation then requires a reliable way to call `Object.create(null)` and to initialize `Function.prototype.rtti`. To achieve this, compiled TS* programs are linked with a library called `boot.js`. This library is intended to be the first piece of JavaScript that runs on a page—§5 discusses how to make a script run first, reliably, before any adversarial script. `boot.js` takes a clean copy of `Object.create` and stores it in an immutable field. Later, the translated TS* code accesses `Object.create` from this field, rather than as `window.Object.create`, which is an attacker-controlled path. We show a fragment of `boot.js` below, simplified slightly to use

[1] Based on http://davidwalsh.name/json-validation.

[2] http://tools.ietf.org/html/draft-zyp-json-schema-03

field names such as `rtti`, instead of the longer names (less likely to clash with source code) used in our implementation.

```
 1 function boot() {
 2   var Clean     = Object.create(null);
 3   Clean.create = Object.create;
 4   Clean.isTag  = function (t,s,x) { ... }; ...
 5   Clean.wrap   = function (t,s,x) { ... };
 6   Object.freeze(Clean);
 7   Object.defineProperty(Function.prototype, "rtti",
 8   {value:null, writable:false, configurable:false});
 9   Object.defineProperty(window, "Clean",
10   {value:Clean, writable:false, configurable:false});
11 };
12 boot(); boot=undefined;
```

The listing above defines a function `boot` that is run once and then discarded. Within the scope of the function, we construct a `Clean` object in which to maintain pristine copies of functions like `Object.create`, on which our translation relies. Lines 4–5 define functions that implement queries and coercions on runtime types and values—we discuss their implementation in detail in §3. Line 6 calls the JavaScript function `freeze`, to make `Clean` immutable. Line 7 initializes `Function.prototype.rtti` and line 9 registers the `Clean` object at the path `window.Clean`—both use JavaScript's `defineProperty`, to define immutable, non-configurable properties.

2.4 Embedding TS★ in JavaScript The `Clean` object in `boot.js` provides trusted core functionality upon which we can build a secure compiler. In this section, we outline the end-to-end embedding within JavaScript of the `Point` example from §2.2. There are a few broad features of the translation that we focus on:

- Adding runtime-type information to every object and function.
- Checking runtime type information in the any fragment.
- Embedding wrappers to safely export/import values.

The listing below shows the translation of the TS★ function `Point` to JavaScript. The translated code is placed within a single enclosing function to introduce a fresh local scope. Without this precaution, TS★ definitions would implicitly leak into the global un-typed object. The type annotations in TS★ are all erased in JavaScript.

```
 1 function () {
 2  var Point = function (x,y) {
 3    var self = Clean.create(null);
 4    self.rtti = ⦇{}⦈;
 5    write(self, "x", x);
 6    write(self, "y", y);
 7    var tmp   = function(d) {write(self, "x", d);}
 8    tmp.rtti  = ⦇number → unit⦈;
 9    write(self,"setX", tmp);
10    return Clean.setTag(⦇{}⦈,⦇point⦈,self);
11  }
12  Point.rtti = ⦇(any,any) → point⦈;
13  var o = Point(0,0);
14  o.setX(17);
15 }()
```

Line 3: the source empty record `{}` is compiled to a new null-prototype object. Line 4: we set the `rtti` field of `self` to the translation of a source empty record type. Lines 5 and 6: we use the macro *write* to add two properties to the `self` object. This macro (defined in §3) checks that the RTTI of the assigned field (if any) is compatible with the RTTI of the assignee. In this case, since the `rtti` field of `self` is just the empty record, it does not constrain the contents of any of its fields, so these assignments succeed and the fields are added to `self`. Line 7: we translate `setX` and, line 8: we tag it with an `rtti` field recording its source type. We then add it to `self` using *write*. Line 10: the call to `Clean.setTag` checks whether `self`, whose static type is represented by ⦇{}⦈, contains a valid representation of a source `point`. For this, it examines the representation of the type ⦇`point`⦈; notice that the type requires three fields, `x`, `y`, and `setX`;

$$
\begin{array}{llcl}
\textit{Value} & v & ::= & x \mid \texttt{true} \mid \texttt{false} \mid \lambda x{:}t.e \mid D\,\bar{v} \\
\textit{Expr.} & e & ::= & v \mid \{\bar{f} = \bar{e}\} \mid e.f \mid e.f = e' \mid e[e'] \mid e[e'] = e'' \\
 & & \mid & \textbf{let}\ x\ =\ e\ \textbf{in}\ e' \mid e\ e' \mid D\,\bar{e} \\
 & & \mid & \textbf{if}\ e\ \textbf{then}\ e'\ \textbf{else}\ e'' \mid q\langle t\rangle e \mid c\langle t\rangle e \\
\textit{Query} & q & ::= & \textit{isTag} \mid \textit{canTag} \mid \textit{canWrap} \\
\textit{Coercion} & c & ::= & \textit{setTag} \mid \textit{wrap} \\
\textit{Type} & t,u & ::= & \mathsf{bool} \mid \mathsf{T} \mid \mathsf{any} \mid \mathsf{un} \mid t \to u \mid \{\bar{f} :^{\bar{a}} \bar{t}\} \\
\textit{Access} & a & ::= & \mathsf{r} \mid \mathsf{w} \\
\textit{Sig.} & S & ::= & . \mid D{:}\bar{t} \to T \mid S, S' \\
\textit{Env.} & \Gamma & ::= & . \mid x{:}t \mid \Gamma, \Gamma'
\end{array}
$$

Figure 1. Formal syntax of TS★

then it checks that the `self` object contains values in those three fields whose RTTIs are compatible with the requested types `number`, `number`, and `number -> unit`, respectively. Once this check succeeds, `setTag` updates the `rtti` field of `self` to ⦇`point`⦈. An invariant of our translation is that the `rtti` field of every object evolves monotonically with respect to the subtyping relation. That is, `self.rtti` was initially ⦇{}⦈ and evolves to ⦇`point`⦈, where `point <: {}`. The RTTI of `self` may evolve further, but it is guaranteed to always remain a subtype of `point`. Line 12: we add an `rtti` field to the `Point`. Finally, lines 13 and 14: we see the translation of a statically typed fragment of TS★. Pleasantly, the translation there is just the identity.

As shown, the translated program does not interact with its context at all. However, the programmer can choose to export certain values to the context by writing `export function Point(x,y){...}` instead. This instructs the compiler to wrap and export `Point` to the context by inserting the following code after Line 14.

```
win.Point = Clean.wrap(⦇(any,any) → point⦈, ⦇un⦈,Point);
```

3. Formalizing TS★

This section formalizes TS★ by presenting its type system and type-directed translation to JavaScript. We describe in particular our runtime representation of types and the JavaScript implementations of `Q.wrap`, `Q.setTag` and related functions that manipulate translated terms and their types. We conclude this section with a detailed comparison of TS★ with prior gradual type systems.

3.1 Syntax Figure 1 presents our source syntax. To aid in the readability of the formalization, we employ compact, λ-calculus style notation, writing for example $\lambda x{:}t.e$ instead of `function(x:t) {return e;}`. We also write $\bar{e}$ for a sequence of expressions $e_1,\ldots,e_n$, $f(\bar{e})$ for the application $f(e_1,\ldots,e_n)$, and so on. Our formal syntax does not cover the un-typed fragment, since its typing and compilation are both trivial, although, of course, our theorems specifically address the composition of compiled TS★ code with arbitrary JavaScript.

Values v include variables x, Booleans, typed λ-abstractions, and data constructors D applied to a sequence of values. For conciseness, we exclude primitives like numbers and strings, since they can in principle be encoded using data constructors. In practice, our implementation supports JavaScript primitives, and so we use them in our examples. This requires some care, however, since some operations on primitive types can be overwritten by the adversary. In such cases, we cache reliable versions of those primitive operations in the `Clean` object built by `boot.js`

In addition to values, expressions e include record literals, projections of static fields, and assignments to static fields. We also include projections of computed fields $e[e']$ and assignment to computed fields $e[e'] = e''$. It is important to note that records, even records of values, are not values, as in JavaScript evaluating a record returns the heap location where the record value is stored. We have `let`-bindings (corresponding to immutable `var` bindings in our concrete syntax); function application; data constructor appli-

$$\boxed{\Gamma \vdash e : t \rightsquigarrow s} \qquad \frac{\Gamma \vdash e : u \rightsquigarrow s \quad S \vdash u <: t}{\Gamma \vdash e : t \rightsquigarrow s}\ \text{(T-SUB)} \qquad \frac{}{\Gamma \vdash x : \Gamma(x) \rightsquigarrow x}\ \text{(T-X)} \qquad \frac{\Gamma \vdash \bar{e} : \bar{t} \rightsquigarrow \bar{s} \qquad \{\bar{f} :^{\bar{a}} \bar{t}\} = t \uplus u \qquad S \vdash \mathsf{unFree}(t)}{\Gamma \vdash \{\overline{f = e}\} : u \rightsquigarrow \mathit{record}(\overline{f{:}s}, u)}\ \text{(T-REC)}$$

$$\frac{S(D) = \bar{t} \to T \quad \Gamma \vdash \bar{e} : \bar{t} \rightsquigarrow \bar{s}}{\Gamma \vdash D\ \bar{e}\ : T\ \rightsquigarrow\ \mathit{data}(D, \bar{s}, T)}\ \text{(T-D)} \qquad \frac{\Gamma, x{:}t \vdash e : t' \rightsquigarrow s}{\Gamma \vdash \lambda x{:}t.e : t \to t' \rightsquigarrow \mathit{fun}(x, e, s, t \to t')}\ \text{(T-LAM)} \qquad \frac{\Gamma \vdash e : u \uplus \{f :^{\mathsf{w}} t\} \rightsquigarrow s \qquad \Gamma \vdash e' : t \rightsquigarrow s'}{\Gamma \vdash e.f = e' : t \rightsquigarrow s.f = s'}\ \text{(T-WR)}$$

$$\frac{\Gamma \vdash e{:}u \uplus \{f :^{a} t\} \rightsquigarrow s}{\Gamma\ \vdash\ e.f\ :\ t\ \rightsquigarrow\ s.f}\ \text{(T-RD)} \qquad \frac{\Gamma \vdash e : t_1 \to t_2 \rightsquigarrow s \qquad \Gamma \vdash e' : t_1 \rightsquigarrow s'}{\Gamma \vdash e\ e' : t_2 \rightsquigarrow s\ s'}\ \text{(T-APP)} \qquad \frac{\Gamma \vdash e : u \rightsquigarrow s \qquad \Gamma, x{:}u \vdash e' : t \rightsquigarrow s'}{\Gamma \vdash \mathbf{let}\ x\ =\ e\ \mathbf{in}\ e' : t \rightsquigarrow (x = s, s')}\ \text{(T-LET)}$$

$$\frac{\Gamma \vdash e : \mathtt{any} \rightsquigarrow s \qquad \forall i. \Gamma \vdash e_i : t \rightsquigarrow s_i}{\Gamma \vdash \mathbf{if}\ e\ \mathbf{then}\ e_1\ \mathbf{else}\ e_2 : t \rightsquigarrow \mathtt{if}\ (s)\ \{s_1\}\ \mathtt{else}\ \{s_2\}}\ \text{(T-IF)} \qquad \frac{\Gamma \vdash e : t' \rightsquigarrow s \qquad t' \sim t}{\Gamma \vdash q\langle t \rangle e : \mathtt{bool} \rightsquigarrow \mathtt{Clean}.q((\!| t' |\!), (\!| t |\!), s)}\ \text{(T-Q)} \qquad \frac{\Gamma \vdash e : t' \rightsquigarrow s \qquad t' \sim t}{\Gamma \vdash c\langle t \rangle e : t \rightsquigarrow \mathtt{Clean}.c((\!| t' |\!), (\!| t |\!), s)}\ \text{(A-C)}$$

$$\frac{\forall i. \Gamma \vdash e_i : \mathtt{any} \rightsquigarrow s_i}{\Gamma \vdash e_1\ e_2 : \mathtt{any} \rightsquigarrow \mathit{apply}(s_1, s_2)}\ \text{(A-APP)} \qquad \frac{\forall i. \Gamma \vdash e_i : \mathtt{any} \rightsquigarrow s_i}{\Gamma \vdash e_1[e_2] : \mathtt{any} \rightsquigarrow \mathit{read}(s_1, s_2)}\ \text{(A-RD)} \qquad \frac{\forall i. \Gamma \vdash e_i : \mathtt{any} \rightsquigarrow s_i}{\Gamma \vdash e_1[e_2] = e_3 : \mathtt{any} \rightsquigarrow \mathit{write}(s_1, s_2, s_3)}\ \text{(A-WR)}$$

$$\boxed{S \vdash t <: t'} \qquad \frac{}{S \vdash t <: t} \qquad \frac{S \vdash t <: t'' \quad S \vdash t'' <: t'}{S \vdash t <: t'} \qquad \frac{S \vdash t' <: t \quad S \vdash u <: u'}{S \vdash t \to u <: t' \to u'} \qquad \frac{S \vdash t <: t'}{S \vdash \{f :^{\mathsf{r}} t\} \uplus u <: \{f :^{\mathsf{r}} t'\} \uplus u} \qquad \frac{S \vdash \mathsf{unFree}(t)}{S \vdash t <: \mathtt{any}} \qquad \frac{S \vdash \mathsf{unFree}(t)}{S \vdash t \uplus u <: u}$$

$$\boxed{S \vdash \mathsf{unFree}(t)} \qquad \frac{}{S \vdash \mathsf{unFree}(\mathtt{bool})} \qquad \frac{}{S \vdash \mathsf{unFree}(\mathtt{any})} \qquad \frac{\forall D{:}\bar{t} \to T \in S[T]. S \vdash \mathsf{unFree}(\bar{t})}{S \vdash \mathsf{unFree}(T)} \qquad \frac{\forall i. S \vdash \mathsf{unFree}(t_i)}{S \vdash \mathsf{unFree}(t_1 \to t_2)} \qquad \frac{S \vdash \mathsf{unFree}(\bar{t})}{S \vdash \mathsf{unFree}(\{\bar{f} :^{\bar{a}} \bar{t}\})}$$

$$\boxed{t \sim t'} \qquad \frac{}{t \sim t} \qquad \frac{t' \sim t}{t \sim t'} \qquad \frac{}{\mathtt{any} \sim t} \qquad \frac{}{\mathtt{un} \sim t} \qquad \frac{t \sim t' \qquad u \sim u'}{t \to t \sim t' \to u'} \qquad \frac{\mathsf{fields}(u_0) \cap \mathsf{fields}(u_1) = \emptyset \qquad \forall j. a_j = a'_j \wedge t_j \sim t'_j}{\{\bar{f} :^{\bar{a}} \bar{t}\} \uplus u_0 \sim \{\bar{f} :^{\bar{a'}} \bar{t'}\} \uplus u_1}$$

where

`let x =` s `in` s'	$\triangleq$	`function(x){return` s'`;}(`s`)`
$\mathit{record}(\overline{f} : \overline{s}, t)$	=	`let x=Clean.create(null) in (x.`$\overline{f}$`=`$\overline{s}$`, x.rtti=`$(\!| t |\!)$`)`
$\mathit{data}(D, \bar{s}, T)$	=	`let x=Clean.create(null) in (x[`$\bar{i}$`]=`$\bar{s}$`, x.c=`$(\!| D |\!)$`, x.rtti=`$(\!| T |\!)$`, x)`
$\mathit{fun}(x, e, s, t \to t')$	=	`let f=function(`x`){var` $\mathit{locals}(e)$`; return` s`;} in (f.rtti=`$(\!| t \to t' |\!)$`, f)`
$\mathit{apply}(s_1, s_2)$	=	`let f=`s_1` in let x=`s_2` in typeof(x)==="function" ? f(Clean.setTag(`$(\!| \mathtt{any} |\!)$`, f.rtti.arg,x)) : Clean.die()`
$\mathit{read}(s_1, s_2)$	=	`let x=`s_1` in let f=`s_2` in typeof(x)==="object" && Clean.hasField(x,f) ? x[f] : Clean.die()`
$\mathit{write}(s_1, s_2, s_3)$	=	`let x=`s_1` in let f=`s_2` in let v=`s_3` in let t = typeof(x)==="object" ? x.rtti : Clean.die() in Clean.mutable(t, f) ? x[f]=Clean.setTag(`$(\!| \mathtt{any} |\!)$`, Clean.hasField(t,f) ? t[f] : Clean.Any(v) : Clean.die()`

$(\!| \mathtt{any} |\!)$	=	`Clean.Any`	$(\!| t \to t' |\!)$	=	`Clean.arrow(`$(\!| t |\!)$`, `$(\!| t' |\!)$`)`	$(\!| \mathtt{un} |\!)$	=	`Clean.Un`
$(\!| T |\!)$	=	`Clean.data("T")`	$(\!| \{\overline{f} :^{\bar{a}} \bar{t}\} |\!)$	=	`let o = Clean.rec()in Clean.addField(o, `$\bar{f}$`, `$(\!| \bar{t} |\!)$`, `$\bar{a}$`==="w")`			

Figure 2. A type-directed translation of TS* to JavaScript

cation; and conditionals. Finally, we have RTTI-based query operations $q\langle t \rangle e$, and coercions $c\langle t \rangle e$.

Types t, u include a number of primitive types (`bool` for boolean values, and `any` and `un` for dynamic values), abstract data types ranged over by T, and records. Record types are written using the shorthand $\{\bar{f} :^{\bar{a}} \bar{t}\}$ to denote the type $\{f_1 :^{a_1} t_1, \ldots, f_n :^{a_n} t_n\}$ where the f_i are distinct and the a_i are accessibility annotations: r for read-only, and w for mutable. We also write $t \uplus u$ for the record type $\{\bar{f_1} :^{\bar{a_1}} \bar{t_1}, \bar{f_2} :^{\bar{a_2}} \bar{t_2}\}$, where $t = \{\bar{f_1} :^{\bar{a_1}} \bar{t_1}\}$ and $u = \{\bar{f_2} :^{\bar{a_2}} \bar{t_2}\}$.

The type system is given with respect to a signature S which maps data constructors D to their type signature, written $\bar{t} \to T$. In places we need to refer to *all* the data constructors for a given abstract data type T in the signature S. We use the shorthand $S[T]$ which is defined as $\{D : \bar{t} \to T \mid S(D) = \bar{t} \to T\}$. We also have a standard type environment Γ binding variables to their types.

Although we have data constructors, pattern matching in TS* is not primitive. Instead, it can be encoded in terms of the other constructs, as defined below. Note that we freely use `&&` and other Boolean operators, as well as `===`, physical equality on TS* values.

$$\mathbf{match}\ e\ \mathbf{with}\ D_{\bar{t} \to T}\bar{x}\ \to e_1\ \mathbf{else}\ e_2 \triangleq$$
$$\mathbf{let}\ y\ =\ e\ \mathbf{in}\ \mathbf{if}\ \left(\mathit{isTag}\langle T \rangle y\ \mathtt{\&\&}\ y.c === \mathtt{"D"}\right)\ \mathbf{then}\ \mathbf{let}\ \bar{x}\ =\ \overline{y[\mathrm{i}]}\ \mathbf{in}\ e_1\ \mathbf{else}\ e_2$$

3.2 Type system and translation Figure 2 defines the judgment $\Gamma \vdash e : t \rightsquigarrow s$, which states that in an environment Γ (along with an implicit signature S), the expression e can be given the type t and be translated to the JavaScript program s. We present the type system declaratively. In practice, our type checker analyzes a fully decorated AST produced by the type inference algorithm of TypeScript, so implementing the system in Figure 2 is straightforward. (A precise description of this type inference algorithm is beyond the scope of this paper.)

At a high level, the type system is designed to enforce the following three properties mentioned in §2:

Static safety TS* programs have no failing dynamic checks during the execution of statically typed sub-terms. We achieve this via two mechanisms: (a) the rules prefixed by (T-) enforce the static typing discipline and they never insert any runtime checks when compiling the program; (b) when a value v:`any` is passed to a context that expects a precise type, e.g. `point`, the compiler inserts instrumentation to ensure that v is indeed at least a `point`. Instrumentation inserted elsewhere in dynamic code also ensures that v henceforth remains at least a `point`. This protects statically typed code from future modifications to v. In the other direction, the type system allows for v :`point` to be passed to `any`-typed context via subtyping.

Dynamic safety The RTTI of $v{:}t$ is always a subtype of t and a sound approximation of v's most precise type—by two mechanisms: (a) v's RTTI initially reflects t and the *setTag* operation ensures that RTTI always evolves towards the more precise types per subtyping, and (b) the rules prefixed by (A-) instrument the translation of the `any`-typed parts of the source to enforce that modifications to v respect its RTTI. (We envision that an IDE can highlight uses of (A-) rules to the programmer as potential failure points.)

Memory isolation un-typed code cannot directly access an object reference that TS* code may dereference; this is enforced by ensuring that the un type is treated abstractly. The only way to manipulate un values is via defensive wrappers, which means that typed code never dereferences an un-typed memory location, and that any-typed references are never be handed directly to the adversary. The subtyping rules are designed to prevent masking the presence of un-values in records using width-subtyping or subtyping to any.

We now turn to describing each of the rules in detail. The first rule in the judgment, (T-Sub), is a subsumption form which shows that a use of subtyping in TS* does not change the translation of a term. The subtyping relation $S \vdash t <: t'$ (also in Figure 2) is mostly standard. Depth subtyping on records is permitted only for immutable fields. The penultimate rule allows all types that do not contain the un type to be a subtype of any. (The auxiliary predicate unFree detects occurrences of un in a type.) Allowing `un <: any` would clearly break our invariants. Allowing `{f:un} <: any` is also problematic, since if a value `v:{f:un}` could be promoted to `v:any`, then `v["f"]` would also have type any, even though it produces an untrusted value. The last subtyping rule provides width-subtyping on records, forgetting the fields to weaken $t \uplus u$ to u, only so long as t contains no occurrences of un.

The rule (T-X) for typing variables is standard. (T-Rec) introduces a record at type u, such that u includes all the un-fields (necessary for compatibility with subtyping). Its compilation allocates a new object, safely sets the fields $\bar{f}$ to $\bar{s}$, and finally adds an `rtti` field containing $(\!|u|\!)$. The rule (T-D) for typing data constructors is similar. The typing of functions with (T-Lam) is standard; however, the translation to JavaScript is a bit subtle: it defines a JavaScript function tagged with an `rtti` field, whose body s is preceded by declarations of all the `let`-bound variables in e, the source function body. These (and other) rules use the JavaScript form $(\bar{e})$, which evaluates every e_i in $\bar{e}$ and returns the last one.

The rules (T-Wr), (T-Rd), and (T-App) are standard. One slight wrinkle in (T-If) is that we rely on JavaScript implicitly converting u to a boolean—this implicit conversion is a safe primitive operation. One could imagine a stricter variant of (T-If) with a runtime check to ensure that u is a boolean, without applying implicit conversions. However, (T-If) as shown is more idiomatic of JavaScript. On the other hand, our treatment of local variables in (T-Let) deviates from JavaScript's `var`-statements—local variables are always immutable in TS*. We make this choice since mutability, if desired, can be encoded by boxing the let-bound value in a mutable record field. In contrast, encoding immutable local variables given only mutable ones is impossible (even with immutable record fields).

The rules (T-Q) and (A-c) cover queries $q\langle t\rangle e$ and coercions $c\langle t\rangle e$. In each case, we have an expression $e{:}t'$ compiled to s, and we apply q or c at type t, so long as t and t' are compatible. Type compatibility is a simple reflexive, symmetric, and *non*-transitive relation, in the spirit of Siek and Taha (2006). We discuss the implementation of these operations in the next subsection.

The remaining (A-) rules instrument the translated programs to ensure safety. In (A-App), we first check that s is a function. Then, before calling the function, we tag the argument with the type of the function's parameter. (A-Rd) simply checks that s_1 is an object and that s_1 has field s_2. (A-Wr) checks that s_1 is an object. It then checks that s_1's RTTI allows for field s_2 to be written. If s_1's RTTI does not contain s_2, it is treated as a new property addition—deleting a property if it is not present in the RTTI is also straightforward, although we do not cover it here. Otherwise, it should contain a mutable s_2 field, and before writing, s_3 is tagged with the type expected by s_1's RTTI.

3.3 Implementing RTTI-based coercions and wrappers As described in §2.3, runtime support for compiled TS* programs is provided by the immutable `Clean` object installed in the global namespace by the first-starter script `boot.js`. This section discusses in detail the `setTag` and `wrap` operations provided by the runtime.

Updating RTTI with setTag The form $\mathit{setTag}\langle t\rangle(e{:}t')$ is compiled to `Clean.setTag(`$(\!|t'|\!)$`,`$(\!|t|\!)$`,`s`)`, where `e` compiles to s. In this case, the first argument $(\!|t'|\!)$ is redundant, since it can be recovered from the RTTI of s (which, by dynamic safety, must be a refinement of t')—we keep the first argument for uniformity with wrappers, as discussed below. The call `Clean.setTag(`$(\!|t'|\!)$`,`$(\!|t|\!)$`,s)` boils down to `setTagAux(s,`$(\!|t|\!)$`,false)`, whose implementation is shown in Figure 3 (on the left). This code tries to update the RTTI of `s` to be some subtype of $(\!|t|\!)$ and fails otherwise. There are two main concerns that the code addresses: (1) the code should not diverge if `s` is a cyclic data structure (unless `die` is called); (2) if the tag of `s` is updated, then `s` must represent a value of type t and the new tag must be a subtype of both the old tag and t. These two concerns are complementary, since certain types (e.g., non-recursive types or types without mutable fields) specifically require all their values to be acyclic. The main idea of `setTagAux(x,t,b)` is to traverse the objects reachable from `x` in a depth-first manner, at each point checking that each field required by t is present in the object at the expected type (lines 11 and 21). Each object encountered in the traversal is temporarily marked (line 7) as visited by assigning `meet(curTag,t)` to `x.tmp`. The assigned value is the greatest subtype of both `curTag` and `t` (which may not exist, in which case the operation fails by calling `die`). Cycles are only permissible when traversing mutable object references and cycles are detected by looking for a marker in `x.tmp`. If the traversal succeeds, the temporary marker is made permanent by updating `x.rtti` (line 4); if it fails, a fatal error is raised by calling `die`, which exhausts the JavaScript stack. This is drastic but effective; friendlier failure modes are feasible too.

Coercions based on `setTag` can be overly conservative, particularly on higher-order values. For example, trying to coerce the identity function `id : any -> any` to the type `bool -> bool` using `setTag` will fail, since `any` $\not<:$ `bool`. As such, RTTI-based coercions are most effective when working with mutable first-order objects. One way to broaden the scope of RTTI-based coercions is, of course, to enrich the type language, e.g., to include polymorphism, or even refinement types—we plan to explore this in the future. Additionally, we provide operations to query RTTI, $\mathit{isTag}\langle t\rangle e$ and $\mathit{canTag}\langle t\rangle e$, that (respectively) allow programmers to query the current tag of e and to test whether or not a `setTag` operation on e would succeed. When RTTI-based coercions are inapplicable (e.g., on un-typed values) or overly conservative (e.g., on functions), the `wrap` form is handy—we describe it next.

Wrappers When $e{:}t'$ compiles to s, $\mathit{wrap}\langle t\rangle(e{:}t')$ compiles to `Clean.wrap(`$(\!|t|\!)$`,`$(\!|t'|\!)$`,`s`)`—Figure 3 shows part of its implementation (on the right). Line 2 is the case where both $(\!|t|\!)$ and $(\!|t'|\!)$ are function types. As expected, we apply a higher-order cast to s, wrapping the argument, applying s, and then wrapping the result. Thus, `wrap(`$(\!|$`any`$\rightarrow$`any`$|\!)$`, `$(\!|$`bool`$\rightarrow$`bool`$|\!)$`, id)` succeeds where the corresponding `setTag` fails.

Securely importing from un The more interesting case of `wrap` is when the source type is `un` and the target type is `any` (line 7), which occurs when importing an un-typed value from the adversary into the typed fragment of TS*. At line 8, we use JavaScript's `typeof` operator to examine the simple type of `x`, the `un`-typed value being coerced. If `x` has a primitive type, then the coercion is just the identity. If `x` is an object, we aim to enumerate all its fields, then wrap and copy each of those fields into a newly allocated `any`-typed object. This requires some care, however, since directly enumerating or accessing the fields of an `un`-typed objects causes callbacks to the `un`-typed adversary, potentially breaking memory isolation by leaking `any`-typed objects on the stack to the context. As a defense, we build on an idea from Fournet et al. (2013), as follows. At line 14, we allocate a new object `r` into which we will wrap and copy the fields of `x`. Next, we build a closure (`stub` at line 15) which captures

```
 1 function setTagAux(x, t, cycleok) {
 2  var curTag = tagOf(x);
 3  if (tmp in x) { cycleok ? curTag = x.tmp : die(); }
 4  function ok(){ x.rtti=x.tmp;delete x.tmp;return x; }
 5  if (subtype(curTag, t)) return x;
 6  if (typeof x !== "object"||x.tmp) die();
 7  x.tmp = meet(curTag, t);
 8  switch (t.t) {
 9   case "record":
10    if (curTag.t!=="record"||curTag.t!=="Any") die();
11     foreach(t, function(fld, ft) {
12      if (!hasField(x,fld)) die();
13      if (hasField(curTag,fld) &&
14          ft.mut!==curTag[fld].mut)) die();
15      setTagAux(x[fld], ft.typ, ft.mut);
16     }); return ok();
17   case "data"   :
18    if(curTag.t !== "Any") die();
19    if(!hasField(x,"c")) die();
20    var ct = constructors(t)[x.c];
21    foreach (ct, function (ix, arg) {
22     if (!hasField(x,ix)) die();
23     setTagAux(x[ix],arg,false);
24    }); return ok();
25   default : die();
26 }}
```

```
 1 function wrap (src, tgt, x) {
 2  if (src.t === "arrow" && tgt.t === "arrow") {
 3   var f = function (y) { return
 4     wrap(src.ret,tgt.ret,x(wrap(tgt.arg,src.arg,y)));
 5   }; f.rtti = tgt; return f;
 6  } ...
 7  else if (src.t === "un" && tgt.t === "any") {
 8   switch (typeof x) {
 9    case "undefined":
10    case "string":
11    case "number":
12    case "boolean": return x;
13    case "object":
14     var r = Clean.create(null);
15     var stub = function () {
16       foreach(x, function (p, v) {
17        r[p] = wrap(⦇un⦈, ⦇any⦈,v);
18      })};
19     callUn(stub); return r;
20    case "function":
21     var un2any=function (x){return wrap(⦇un⦈,⦇any⦈,x);}
22     var any2un=function (x){return wrap(⦇any⦈,⦇un⦈,x);}
23     var f = upfun(any2un,un2any)(x);
24     f.rtti = ⦇any → any⦈; return f;
25    default: die();
26 }}}
```

Figure 3. Implementations of `Clean.setTag(u,t,x)` $\triangleq$ `setTagAux(x,t,false)` and `Clean.wrap` (selected fragments)

`r` and `x`, but otherwise takes no arguments and returns nothing. As such, the `stub` can be viewed as a function from `un` to `un`. In the body of `stub`, we enumerate the fields of `x`, then wrap and copy each of them into `r`. All that remains is to call `stub` and return `r`. The call itself is done using `callUn`, which provides a safe way to call an `un`-typed function without breaking memory isolation. The function `callUn(f)` is defined as `upfun(id,id)(f)(undefined)`, where `id` is the identity function, and `upfun` is the wrapper for importing `un`-typed functions from Fournet et al. (2013) (see Figure 4 in that paper). When `x` is a function (line 20), we again use `upfun`, this time exporting the argument and then importing the result back—the returned value has type `any -> any`.

***Securely exporting to* un** The implementation of `wrap(⦇t⦈,⦇un⦈)(s)` for $t \neq$ un, exports `s` to the context. This case is somewhat simpler, since `s` is a typed value with RTTI which can be safely probed. If s is a primitive, it can be exported as is. If `s` is an object, `wrap` creates an empty object `x`, enumerates all the fields `f` of `s`, exports them recursively and adds them to `x`. For functions, it uses Fournet et. al.'s `downfun` to export it at type `un -> un`.

3.4 Discussion and related work on gradual typing

Languages that mix static and dynamic types date back at least to Abadi et al. (1991) and Bracha and Griswold (1993). Gradual typing is a technique first proposed by Siek and Taha (2006), initially for a functional language with references, and subsequently for languages with various other features including objects. Several others have worked in this space. For example, Flanagan (2006) introduces hybrid typing, mixing static, dynamic and refinement types; Wadler and Findler (2009) add *blame* to a gradual type system; Herman et al. (2010) present gradual typing with space-efficient wrappers; Bierman et al. (2010) describe type dynamic in C$^\sharp$; and Ina and Igarashi (2011) add gradual typing to generic Java.

Our system is distinct from all others in that it is the first to consider gradual typing for a language *embedded* within a larger, potentially adversarial environment via the type un. We are also, to the best of our knowledge, the first to consider gradual typing as a means of achieving security.

To compare more closely with other systems, let us set aside un for the moment, and focus on the interaction between any and statically typed TS*. Previous type systems mediate the interactions between static- and any-typed code by *implicitly* attaching casts to values. Higher order casts may fail at a program point far from the point where it was inserted. To account for such failures, blame calculi identify the cast (with a label to indicate the term or context) that causes the failure—Siek et al. (2009) survey blame calculi based on the errors they detect, points of failures, and casts they blame. In contrast, Interactions between static- and any-typed TS*is based primarily on RTTI. Casts (wrappers in our terminology) are never inserted implicitly, although they are made available to the programmer. This design has the following advantages.

Preservation of object identity Object identity in JavaScript is a commonly used feature. Since TS* does not *implicitly* attach casts to values, it never implicitly breaks object identity in the source during compilation. Previous gradual type systems with implicit casts would always break object identity.

Space efficiency Casts can pile up around a value, making the program inefficient. Herman et al. (2010) introduce a novel cast-reduction semantics to gain space-efficiency. Our approach is also space efficient (there is only one RTTI per object) but does not require cast-reduction machinery.

Static safety and eager failures In contrast to our RTTI-based mechanism, statically typed code in other gradual type systems could fail (although blame would help them ascribe it to the any-typed code). Consider the following TS* example.

```
let v:any = {f:true} in (λr:{f:ʷint}. r.f) v
```

Compiling this term using (A-APP) introduces a `setTag` on the argument `v` at type `{f:ʷint}`. The `setTag` operation, at runtime, recursively checks that `v` is a `{f:ʷint}`, and expectedly fails. Thus, the failure happens prior to the application, a failure strategy called *eager* in prior works. Herman et al. (2010) also argue that their system can provide eager failures, but transposed to their notation (with the record replaced by a `ref any`), the failure occurs at the property read within the statically typed λ-term, breaking static safety. When eager runtime checking seems too strict, TS* wrappers provide an escape hatch. Arguably, for our security applications, a predictably uniform eager-failure strategy is a suitable default.

Dynamic safety and blame With no failures inside statically typed code to explain at runtime, blame seems less useful with RTTI-based coercions. However, because we enforce dynamic safety

(RTTI evolves monotonically), failures may now arise in any-typed code, as in the following example.

```
let v:any={f:true} in (λr:{f:ʷbool}.r.f) v; v.f="hi"
```

This time, the setTag of v to {f:^{w}int} succeeds, and it modifies v's RTTI to be {f:^{w}int}. But now, the update of v.f to "hi" fails. This failure in the any-typed fragment should be blamed on the setTag operation instrumented at the application. We plan to pursue the details of this seemingly new notion of blame as future work.

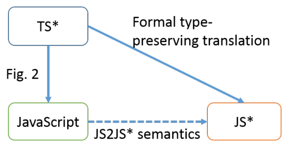
Gradual typing for monotonic objects Independently, in an unpublished manuscript, Siek et al. (2013) investigate gradual typing for a Python-like language (based on an earlier abstract presented at the STOP workshop in 2012). Like us, they notice that constraining the type of a mutable object to evolve monotonically at runtime enables statically typed code to be compiled without runtime checks, leading to improved performance and safety for that fragment. There are some important differences, however. First, Siek et al. require monotonicity with respect to an intentionally naïve subtyping relation. For example, (transposed to our notation) the type of an object can evolve from t ={f :^{w} any} to t' ={f :^{w} number}, although t' is not a subtype of t. This means that, in their system, writing `o.f = true` can fail at runtime, even if o has static type t, which is prevented by TS*. This difference stems perhaps from differing perspectives on static safety: Siek et al. are willing to tolerate runtime errors in code whose typing somewhere employs the type any; in our system, runtime errors may occur only when any-typed values are eliminated (i.e., in our (A-) rules only). Syntactically localizing errors in this manner does not appear as straightforward with naïve subtyping. Additionally, Siek et al. do not aim to model an adversarial context—there is a single dynamic type, not two (i.e., only any, no un). Exploring how to adapt our type un to defend against adversarial Python contexts would be an interesting line of future work.

4. Metatheory

This section formally establishes memory isolation, static safety, and dynamic safety for TS* programs translated to JavaScript. Clearly, such a proof requires a formal semantics for JavaScript—we rely on JS*, a translation semantics for JavaScript developed by Swamy et al. (2013), which is in turn based on λJS (Guha et al. 2010). We provide a brief review of JS*, define the central invariants of our translation, and describe our main theorem. A full formalism, including all proofs, is available online.

4.1 A review of JS* and our high-level proof strategy JS* is a subset of F* (Swamy et al. 2011), a programming language whose type system allows expressing functional-correctness properties of higher-order effectful programs. Prior work defines a semantics for JavaScript via translation to JS*. The semantics is implemented by a tool, called JS2JS*, that translates JavaScript concrete syntax to JS*. This semantics for JavaScript has been used previously both as a means of verifying JavaScript source programs (after translation to JS*) as well as in Fournet et al.'s proof of full abstraction from f* to JavaScript. At its core, JS* provides a mechanically verified library called *JSVerify* that tries to faithfully model most security-relevant details of JavaScript, including, for example, its object model and its calling convention. The metatheory of TS* is stated in terms of its translation to JS*, i.e., programs that can be type-checked against the *JSVerify* API. The validity of our theorem depends on JS* being a faithful model of JavaScript, an assumption that can be checked separately, e.g., by semantics testing.

To set up the formal machinery, we develop a model of our compiler by transposing the translation judgment in Figure 2 to instead generate JS* code. The relationship among these translations is depicted alongside. The translation from TS* to JS* can be seen as the composition of the translation from TS* to JavaScript, and then from JavaScript to JS*. Our main theorem is stated as a type-preservation result from TS* to JS*, where the types in JS* are precise enough to capture our desired invariants, i.e., static safety, dynamic safety, and memory isolation.



Monadic computation types with heap invariants All computations in js* are typed in a state monad of predicate transformers, *iDST*, which is parametrized by a heap-invariant predicate *HeapInv* and a heap-evolution predicate δ. The type *iDST a WP* is the type of a computation, which for any post-condition *post*, when run in an initial heap h, may diverge or else produce a result v:a and a final heap h' that satisfy the formula $post\ v\ h' \wedge HeapInv\ h' \wedge \delta\ h\ h'$, so long as $HeapInv\ h \wedge WP\ post\ h$ is valid. Additionally, all intermediate heaps in the computation satisfy *HeapInv*, and every intermediate heap is related to all its successors by δ. That is, in *iDST a WP*, *a* is the result type, and *WP* is a predicate transformer that computes a pre-condition for the computation with respect to *post*, any desired post-condition; *HeapInv* is an invariant on a heap, and δ is a reflexive and transitive relation constraining how the heap evolves. Our online materials also account for exceptions and fatal errors; however, we gloss over them here for lack of space.

The main idea behind our proof is that a TS* term e:t is translated to a JS* term e' of type *iDST dyn* $WP_{[\![t]\!]}$, where

$$WP_{[\![t]\!]} = \Lambda post.\lambda h.\ Ok\ (loc(e))\ h \wedge \forall v\ h'.\ [\![t]\!]\ v\ h' \Longrightarrow post\ v\ h'$$

This ensures that, if e' is run in an initial heap h satisfying $HeapInv\ h$ and $Ok\ (loc(e))\ h$ (meaning that all free-variables of the source-term e are correctly promoted to the heap in JS*), then either (1) it will terminate with a result v and a final heap h' satisfying $[\![t]\!]\ v\ h'$; or (2) it will diverge. The code may also raise an exception or throw a fatal error, but all the while during the execution of e', *HeapInv* will be true, and the heap will evolve according to δ. This result holds even when e' is linked with arbitrary JavaScript code—adapting the universal typability lemma of Fournet et al. (2013), JavaScript code can always be typed in JS* at a type corresponding to un.

Our main task then is to carefully define *HeapInv*, δ and $[\![t]\!]$ such that they capture our desired invariants, and then to prove that translated programs are well-typed in JS*.

4.2 Invariants of the translation To prove memory isolation, JS* provides a partitioned heap model. Every object reference l:loc carries a tag, $l.tag$, which records the name of the compartment into which the reference points, i.e., each compartment is a disjoint fragment of the domain of the heap. There are six compartments in the heap model. The *Ref* compartment holds objects corresponding to TS* records, datatypes, and RTTI; the *Abs* compartment holds function objects; the *Clean* compartment holds the Clean object initialized and then frozen by boot.js; and the *Un* compartment belongs to the adversary. We focus primarily on properties of these first four compartments. The remaining two compartments, *Inv* and *Stub*, are inherited unchanged from Fournet et al. (2013)—the former is for maintaining local variables, and the latter for tracking function objects used to make safe callbacks to the attacker.

Refined type dynamic All JavaScript values (including the translation of TS*) are represented as JS* values of type *dyn*, defined below. We show only three representative cases. d=$Str\ s$ is an injection of s:$string$ into type *dyn*, where the refinement $TypeOf\ d$=$string$ recalls the static type. For object references $Obj\ (l$:$loc)$, the refinement is l's tag, i.e., {*Ref*, *Abs*, *Un*, *Clean*}. Finally, for functions, *Fun o f* builds a value of type *dyn* from a function closure f and the JavaScript object o for that closure. Its refinement is the predicate transformer of f.

$$\begin{array}{l}\textbf{type}\ \mathit{dyn} = \ldots\\ \mid \mathit{Str}\colon \mathit{string} \to d{:}\mathit{dyn}\{\mathit{TypeOf}\ d{=}\mathit{string}\}\\ \mid \mathit{Obj}\colon l{:}\mathit{loc} \to d{:}\mathit{dyn}\{\mathit{TypeOf}\ d{=}l.\mathit{tag}\}\\ \mid \mathit{Fun}\colon \forall \mathit{WP}.\ o{:}\mathit{dyn} \to (\mathit{this}{:}\mathit{dyn} \to \mathit{args}{:}\mathit{dyn} \to \mathit{iDST}\ \mathit{dyn}\ (\mathit{WP}\ o\ \mathit{args}\ \mathit{this}))\\ \to d{:}\mathit{dyn}\{\mathit{TypeOf}\ d{=}\mathit{WP}\}\end{array}$$

Translation of types To recover the precision of TS* types in JS*, we translate source types to predicates on *dyn*-typed values and the heap: $\llbracket t \rrbracket\ d\ h$ states that the value *d:dyn* is the translation of a source value of type t in the heap h. The translation of a type is with respect to a heap, since a source value allocated at the type $\{f :^a \mathit{number}\}$, may evolve to become a value of type $\{f :^a \mathit{number}, g :^a \mathit{number}\}$ in some subsequent heap. This is in contrast to, and a significant generalization of, the translation of f* to JS*, where a value's type does not evolve and is not subject to subtyping.

$$\begin{array}{lcl}
\llbracket \mathtt{string} \rrbracket\ d\ h & = & \mathit{TypeOf}\ d{=}\mathit{string}\\
\llbracket \mathtt{un} \rrbracket\ d\ h & = & \mathit{IsUn}\ d\\
\llbracket t \rrbracket\ d\ h & = & \exists u <: t.\ \mathit{Tagged}\ u\ d\ h \quad \text{if } t \not\in \{\mathtt{string,\ un}\}\\
\\
\mathit{Tagged}\ u\ d\ h & = & \mathtt{"rtti"} {\in}\ \mathit{dom}\ h[d] \wedge \mathit{Rep}\ u\ (h[d][\mathtt{"rtti"}])\ h \wedge \mathit{Is}\ u\ d\ h\\
\\
\mathit{Is}\ \mathtt{any}\ d\ h & = & \mathit{Prim}\ d \vee \mathit{Fun\_is}\ d \vee (\mathit{TypeOf}\ d{=}\mathit{Ref} \wedge \mathit{restAny}\ \{\}\ d\ h)\\
\mathit{Is}\ T\ d\ h & = & \mathit{TypeOf}\ d{=}\mathit{Ref} \wedge \bigvee_{D_{\bar{t}\to T}}\ h[d][\mathtt{"c"}]{=}\mathit{Str}\ \mathtt{"D"}\\
 & \wedge & \wedge_i\ \llbracket t_i \rrbracket\ (h[d][i])\ h \wedge \mathit{restAny}\ \{1..n\}\ d\ h\\
\mathit{Is}\ \{\bar{f} :^{\bar{a}} \bar{t}\}\ d\ h & = & \mathit{TypeOf}\ d{=}\mathit{Ref} \wedge \bar{f} \subseteq \mathit{dom}\ h[d]\\
 & \wedge & \llbracket \bar{t} \rrbracket\ h[d][\bar{f}]\ h \wedge \mathit{restAny}\ \bar{f}\ d\ h\\
\mathit{Is}\ (t_1 \to t_2)\ d\ h & = & \mathit{TypeOf}\ d = \lambda o\ \mathit{args}\ \mathit{this}.\Lambda p.\lambda h.\\
 & & \quad \llbracket t_1 \rrbracket\ h[\mathit{args}][\mathtt{"0"}]\ h \wedge \forall r\ h'.\ \llbracket t_2 \rrbracket\ r\ h' \Longrightarrow p\ r\ h'\\
\mathit{restAny}\ \mathit{fs}\ d\ h & = & \forall f {\in} \mathit{dom}(h[d]) \backslash \mathit{fs}.\ \neg \mathit{Reserved}\ f \Longrightarrow \llbracket \mathtt{any} \rrbracket\ h[d][f]\ h
\end{array}$$

Since strings are immutable $\llbracket \mathtt{string} \rrbracket\ d\ h$ does not depend on h. Likewise, an un-typed value always remains un-typed—we define *IsUn* shortly. For other types, $\llbracket t \rrbracket\ d\ h$ captures the subtyping relation, stating that there exists a type $u <: t$ such that the value's rtti is tagged with the runtime representation of u (the predicate $\mathit{Rep}\ u\ (h[d][\mathtt{"rtti"}])\ h$), and $\mathit{Is}\ u\ d\ h$, i.e., the value d can be typed at u in h. $\mathit{Is}\ \mathtt{any}\ d\ h$ states that d is either a primitive (e.g., a string), a function, or a location in the *Ref* heap where all its non-reserved fields (excluding, e.g., "rtti") are typeable at any (*restAny*). For datatypes and records, we require d to be a location in the *Ref* heap, with the fields typed as expected, and with all other fields not mentioned in the type being any. The case for functions is most interesting. $\mathit{Is}\ (t_1 \to t_2)\ d\ h$ states that d's predicate transformer builds a pre-condition that requires the first argument to satisfy $\llbracket t_1 \rrbracket$. (All JavaScript functions are variable arity, receiving their arguments in an array; however a non-un-typed TS* function will only read the first one.) In return, the predicate transformer ensures that the result r (if any) will satisfy $\llbracket t_2 \rrbracket$ (recall that we are ignoring exceptions here).

Un values The predicate $\mathit{IsUn}\ v$ defines when the value v could have been produced by, can be given to, or is accessible by the context. *Un* values include primitives; references to objects in the *Un* heap; or the immutable *Clean* object (which is reachable from the global object). Additionally, *Un* values can be functions whose specification indicates that it takes *Un* arguments to *Un* results.

$$\begin{array}{l}\mathit{IsUn}\ x \triangleq \mathit{TypeOf}\ x \in \{\mathit{bool}, \mathit{string}, \mathit{float}, \mathit{Un}, \mathit{Clean}\} \vee \mathit{TypeOf}\ x{=}\mathit{Un2Un}\\ \mathit{Un2Un} \triangleq \lambda o\ \mathit{args}\ \mathit{this}\ \mathit{post}\ h.\ \mathit{IsUn}\ o \wedge \mathit{IsUn}\ \mathit{this} \wedge \mathit{IsUn}\ \mathit{args}\\ \wedge\ (\forall\ r\ h'.\ \mathit{IsUn}\ r \Longrightarrow \mathit{post}\ r\ h')\end{array}$$

HeapInv, the global heap invariant Our main heap invariant is a property of every location x in the heap. Its full definition contains seven clauses; we show the most important ones.

$$\begin{array}{lll}
\mathit{HeapInv}\ h \triangleq & \multicolumn{2}{l}{\forall x.\ x \in \mathit{dom}\ h \Longrightarrow}\\
\quad (1) & & (x.\mathit{tag}{=}\mathit{Un} \Longrightarrow \forall x \in \mathit{dom}\ h[x].\ \mathit{IsUn}\ h[x])\\
\quad (2) & \wedge & (x.\mathit{tag}{\in}\{\mathit{Ref}, \mathit{Abs}\} \wedge \mathtt{"rtti"}{\in}\mathit{dom}\ h[x] \Longrightarrow \exists t.\ \mathit{Tagged}\ t\ d\ h)\\
\quad (3) & \wedge & (x.\mathit{tag}{=}\mathit{Ref} \Longrightarrow \mathit{TypeOf}\ h[x][\mathtt{"@proto"}]{=}\mathit{Null})\\
\quad (4) & \wedge & (x.\mathit{tag}{=}\mathit{Clean} \Longrightarrow \mathit{CleanSpec}\ h[x]\ h)
\end{array}$$

Clause (1) asserts that all the contents of an *Un* object are also *Un*. Clause (2) asserts that every object in the *Ref* and *Abs* compartment with an "rtti" field is tagged properly. Clause (3) additionally specifies that every object in the *Ref* heap has a null prototype. Clause (4) asserts that the *Clean* object is specified by *CleanSpec*, which gives a type to each of its fields.

Within this invariant are two key properties of TS*. The first clause guarantees that the only values reachable from a location in the *Un* heap are themselves un-values. Therein lies our memory isolation property—the adversary can never meddle with TS* objects directly, since these reside in the *Ref* and *Abs* heap, which are disjoint from *Un*. The invariant, in its second clause, also captures dynamic safety, i.e., every object in the *Ref* and *Abs* heap, once tagged with RTTI are properly typed according to it.

δ, the heap evolution invariant The full definition of δ has 4 clauses; we show the main one below: δ ensures that, for all objects in the *Ref* and *Abs* heaps, their "rtti" fields only evolve monotonically downward, according to the subtyping hierarchy.

$$\begin{array}{l}\delta\ h0\ h1 \triangleq \forall l \in \mathit{dom}\ h0, t_0, t_1.\\ l.\mathit{tag} \in \{\mathit{Ref}, \mathit{Abs}\} \wedge \mathit{Rep}\ t_0\ h0[l][\mathtt{"rtti"}]\ h0 \wedge \mathit{Rep}\ t_1\ h1[l][\mathtt{"rtti"}]\ h1\\ \Longrightarrow t_1 <: t_0\end{array}$$

A relatively easy lemma derivable from these definitions implies our static safety property. In particular, Lemma 1 guarantees that if a value v (potentially a location to a heap-resident record) is typeable at type t in some initial heap h_0 then as the program's heap evolves according to δ, v remains typeable at t. This ensures that it is safe for TS* to ascribe a value a static type, since that type is an invariant of the value at runtime.

Lemma 1 (Static safety: δ preserves the interpretation of types)**.** *For all values v, heaps h_0 and h_1 such that HeapInv h_0, HeapInv h_1 and $\delta\ h_0\ h_1$, if for some t we have $\llbracket t \rrbracket\ v\ h_0$ then $\llbracket t \rrbracket\ v\ h_1$.*

Finally, our main theorem, as promised, is a type preservation result that guarantees memory isolation, dynamic safety and static safety for TS* programs translated to JS*. In the hypothesis, the relation $\Gamma \vdash_f e : t \rightsquigarrow e'$ is the formal translation of TS* to JS* (the index f is the name of the current function object in e'; a technical detail). The judgment in the conclusion of the theorem asserts that, in a translated environment, e' has a type that is described by the predicate transformer corresponding to the source type t. As explained in §4.1, this ensures that the translated program respects the heap invariants and, if it terminates normally, produces a t-typed result.

Theorem 1 (Type preservation)**.**
Given a TS* context Γ, an expression e and a type t,
if $\Gamma \vdash_f e : t \rightsquigarrow e'$ for some JS* expression e' and function object f,
then $\llbracket \Gamma \rrbracket, f{:}\mathit{dyn}, \Gamma_{loc(e)} \vdash e' : \mathit{iDST}\ \mathit{dyn}\ \mathit{WP}_{\llbracket t \rrbracket}$.

We conclude our formal development with a few remarks on the scope of our theorem and the style of its proof. First, our result is applicable to the translation of TS* to JavaScript only inasmuch as JS* is an accurate model of all of JavaScript—Fournet et al. (2013) argue for how JS* is adequate for all security-relevant features of JavaScript. Regardless, the availability of these semantics together with its program logic is what made our proof feasible. Given an operational semantics, but lacking a program logic, our proof would have been mired in tedious inductions over the operational semantics. With JS*, we were able to carry out our proof as an induction over the compilation relation, and use the type system of JS* to structure and prove our invariants.

Second, our result includes within its trusted computing base the correspondence of `boot.js` to the *CleanSpec* predicate in the last clause of *HeapInv*. While it is perhaps standard for compiler and verification projects to rely on a small amount of trusted code, we would like to do better. In particular, we aim to use the JavaScript verification toolchain developed by Swamy et al. (2013) to verify `boot.js` for compliance with *CleanSpec*—at the time of writing, this was still incomplete. More substantially, we would also like to build a translation validation pipeline for our compiler implementation, reflecting the generated JavaScript back into JS* for verification, i.e., we would like our compiler implementation to be also formally type-preserving.

5. Securely deploying TS* programs

The guarantees of TS* depend on `boot.js` being the first script to run on a web page. Many prior works have implicitly assumed that scripts are always executed in the order in which they appear on the page (Jim et al. 2007; Magazinius et al. 2010; Taly et al. 2011), but, as we explain, this is a naïve view. Instead, we develop a standards-based mechanism that ensures that our scripts run first.

5.1 Securely bootstrapping the TS* runtime Suppose a script s is lexically the first element in the header of a page located at a URL `U = http://W.com/page.html`; one may expect that it will be guaranteed to run first on any window loaded from `U`. However, this intuition is correct only if the page has not been loaded programmatically from JavaScript by another web page, e.g., within another frame. On a page loaded initially from `U`, the script s will indeed run first. Still, a malicious script running later on the page, or on a different page with the same origin `http://W.com`, may open a window or frame at `U`, and modify all the essential primitives before s begins to run inside the new window frame. This execution order is consistent with the HTML standard (Berjon et al. 2013) and we have confirmed it experimentally on all mainstream browsers.

Hence, any simple first-starter implementation that relies on lexical ordering of script elements will fail if other scripts on the same origin are allowed to open windows or frames. Indeed, the web browser only provides security guarantees at the granularity of an origin (Barth 2011); finer-grained privilege separation between good and bad scripts within the same origin require application-level mechanisms, such as restricting all untrusted scripts to a sub-language like SES$_{light}$ (Taly et al. 2011); loading them in sandboxed iframes with few privileges (Akhawe et al. 2012); or modifying the browser to allow the first-starter script to certify all other scripts running on the page (Jim et al. 2007).

Rather than restrict the functionality of untrusted scripts, we propose a mechanism that ensures that our scripts run first. For a given website, we use two distinct origins:

- `http://W.com`, used primarily as the service origin; it does not serve any resource.
- `http://start.W.com`, that serves HTML pages, including scripts compiled from TS*, but where the first two `<script>` elements on every page are as follows:

```
<script src="http://start.W.com/boot.js"></script>
<script>document.domain = "W.com"</script>
```

The crucial step here is that, after `boot.js` has loaded, the page sets `document.domain` to the parent domain `W.com`. This is a standards-based mechanism (Berjon et al. 2013, 5.3.1) by which the page voluntarily gives up its rights to the `http://start.W.com/` origin for client-side same-origin access across frames and windows. Instead, it adopts an *effective script origin* of `http://W.com`.

All subsequent scripts on the page are unrestricted except that they can only read or write into frames or windows that have an effective script origin of `http://W.com`, and hence they cannot tamper with pages on `http://start.W.com`, even if such pages are loaded programmatically into other frames or windows. In all other ways, their functionality is unimpeded, without the need for expensive translations or messaging protocols, as in previous approaches.

More generally, by placing other trusted scripts after `boot.js` and before the assignment to `document.domain`, we can run scripts that grab reliable copies of builtin libraries, such as JSON and `XMLHttpRequest`, for use by subsequent code.

5.2 Loading scripts with embedded secrets Our first-starter protocol reliably allows `boot.js` to build a trustworthy environment for our compiled scripts. Conversely, we sometimes need a way for scripts to be able to verify that their environment is trustworthy. This is particularly important when compiled scripts contain secret tokens embedded within them, e.g., to authenticate themselves to other servers. Embedding secrets as constants within program text may seem like an elementary mistake, but this is the predominant way of distributing these tokens in a JavaScript setting. Secrets within scripts must first be protected from malicious websites that may try to load our scripts, and second from malicious scripts on our own website `W.com`. In this threat model, many simple countermeasures one may think of are inadequate.

Even if we require a login cookie for authentication before serving the script, a malicious website that an innocent user visits when logged into `W.com` can make a cross-site script request and obtain the script (an attack sometimes called JavaScript hijacking). If we inline the scripts into our page, malicious scripts can read their source code and obtain the token. Even if they are not inlined but served from `http://start.W.com`, malicious scripts can perform an `XMLHttpRequest` to obtain their source code and then read them. Indeed, these are all methods commonly used by cross-site scripting attacks (e.g., the Samy worm) to break token-based security protections on the web.

To protect our scripts from same-origin attackers, we use a third distinct origin to serve our scripts:

- `https://src.W.com`, the secure source server, only serves GET requests for scripts that may embed secret tokens to be shared between the server and the script. The server refuses cross-origin requests and returns scripts as `text/javascript`, so attackers on `https://start.W.com` can execute these scripts but not read its source, due to the same-origin policy. We recommend the use of HTTPS for any pages that contain secrets, but HTTP is adequate if we exclude network adversaries from our threat model.

To protect our scripts against other websites, we need an additional check. Every script served from `src.W.com` is prefixed by a condition on the current webpage location, that is, before making any use of its secret token, the script checks that `window.location.href` actually begins with `http://start.W.com/`. This ensures that the script has a reliable `Clean` object on that page, introduced by `boot.js`.

Experimentally, we found that checking the current location of a script is quite error-prone. Some scripts try to read `document.domain` (see e.g., OWASP CSRFGuard in §6.1) or `document.location`; others rely on `window.location.href` but then use regular expression or string matching to check it against a target origin. All these techniques lead to attacks because a malicious website could have tampered with its `document` object or with the regular expression libraries. We found and reported such attacks to vendors.

Notably, many browsers allow properties like `document.domain` and `window.location.origin` to be overwritten. Our origin check relies on the `window.location.href` object which is specified as unforgeable in the HTML specification (Berjon et al. 2013, 5.2). In practice, however, we found that some browsers incorrectly allow even supposedly unforgeable objects like `window.document` and `window.location` to be shadowed. We have reported these bugs to various browsers and are in discussions about fixes. If the unforgeability of `window.location.href` turns out to be too strong an assumption, we advocate the use of the origin authentication protocol of Bhargavan et al. (2013).

6. Secure web programming with TS*

We have evaluated our compiler by gradually migrating JavaScript sources to TS*, while ensuring that the migrated code (after compilation) exports the same API as the original. We have also written from scratch various utilities like a JSON parser and a reference monitor for HTML5 localStorage. All the code is available online.

We have yet to conduct a thorough performance analysis of our compiler, and to implement further optimizations. But, as mentioned previously, statically typed TS* should incur little, if any, runtime overhead, while un-typed code is unchanged. Understanding and optimizing the performance profile of any-typed code is left as future work.

In this section, we describe how TS* can be used to gradually secure existing access control patterns, as deployed in popular JavaScript APIs. We focus on client-sided, language-based security, relegating most other details to the online materials (notably a description of same-origin policies, protocol- and browser specific security assumptions, and current attacks against them; see also e.g., Bhargavan et al. 2013).

6.1 OWASP CSRFGuard

We first consider the task of securing client code that performs an `XMLHttpRequest` to the page's server.

Cross-Site Request Forgeries (CSRF) Suppose a website `W` has an API available from JavaScript. Authenticating and authorizing access using cookies ensures that only requests from a logged-in user's browser are accepted. Conversely, if the user has any another websites open in her browser, their scripts also get access to the API, and can thus steal or tamper with the user's data on `W`. Such request forgery attacks are persistently listed in OWASP Top 10 vulnerabilities, and have a serious impact on a variety of websites.

CSRF Tokens As a countermeasure to CSRF, a website `W` can inject a fresh, session-specific, random token into every page it serves, and only accept requests that include this token. Other websites cannot see `W`'s pages (thanks to the same origin policy) hence cannot forge requests. Additionally, cookies and tokens can be protected while in transit by using the HTTPS protocol. OWASP CSRFGuard 3 is the most widely-referenced CSRF protection library. As an advanced feature, it provides a token-injection script that transparently protects AJAX calls by intercepting any `XMLHttpRequest`. (Similar protections exist for frameworks like Django and Ruby-On-Rails.) Crucially, the token must be kept secret from other websites, and also from other scripts loaded on the page; otherwise those scripts may use it to perform arbitrary requests directly, or leak it to some other website.

Attacks The original, unwrapped version of their code relies on `window.location`, `String.startsWith`, and `XMLHttpRequest`, which can be tampered with by a malicious script. We found several such attacks where a malicious website could load the OWASP CSRFGuard script, forge the location, and trick it into releasing the token; we are in discussion with the author towards more robust designs, such as the one proposed here for TS*.

CSRFGuard in TS* Following the approach of §2, we migrate to TS* the OWASP proxy that uses the token to provide authentication RPCs to the rest of the library. This small script is typed in TS*, guaranteeing that no malicious script that runs alongside (including the rest of the library) can tamper with its execution.

The TS*proxy, listed below, takes three string arguments: the target URL, the API function name, and the JSON-formatted arguments. It checks that the URL is well-formed and belongs to the current site (to avoid leaking the token to any other site), then it serializes the request as a query string, attaches the token, and makes an AJAX call. Once wrapped, it exports the same interface as before to any (untrusted) scripts loaded on the page. Additionally, it could be directly used by further TS* code.

```
var csrfToken: string = "%GENERATED_TOKEN%"
var targetOrigin: string = "%TARGET_ORIGIN%"
function Rpc(url:string,apifun:string,args:string): any {
  if (String.startsWith(url,String.concat(targetOrigin,"/")) &&
      String.noQueryHash(url)) {
    var m = {method:apifun, args:args, token: csrfToken};
    var request = String.concat("?",QueryString.encode(m));
    var response = xhrGet(String.concat(url,request));}
    return QueryString.decode(response);
  }
  else return "unauthorized URL";
}
```

The first two lines define string literals, inlined by the server as it generates the script—the TS* compilation process ensures, via lexical scoping, that these two strings are private to this script. The `Rpc` function is our secure replacement for `xhrGet`, which performs the actual `XMLHttpRequest`. Compared with the original JavaScript, it includes a few type annotations, and uses either safe copies of builtin libraries, such as `xhrGet` and `String`, or typed TS* libraries, such as `QueryString` (outlined below). Relying on memory isolation and secure loading from TS*, a simple (informal) security review of this script lets us conclude that it does not leak `csrfToken`.

Experimentally, we modified the OWASP library, to isolate and protect the few scripts that directly use the token (such as the proxy above) from the rest of the code, which deals with complex formatting and browser-specific extensions, and is kept unchanged and untrusted. The modified library retains its original interface and functionality, with stronger security guarantees, based on strict, type-based isolation of the token. Its code and sample client- and server-side code are available online. To our knowledge, it is the first CSRF library that provides protection from untrusted scripts.

6.2 Facebook API

Taking advantage of Cross-Origin Resource Sharing (CORS), Facebook provides a client-side JavaScript API, so that trusted websites may access their personal data—once the user has opted in. Interestingly, Facebook also provides a "debug-mode" library, with systematic dynamic typechecks somewhat similar to those automated by TS*, to help programmers catch client-side errors. We focus on two aspects of their large API: the encoding of strings, and cross-domain messaging.

QueryString encoding We give a TS* implementation of the `QueryString` module (mentioned above) for the REST message format used in the Facebook API.

```
function decode (s:string) : any {
 var res = {};
 if (s === "") return res;
 else {
  var params = String.split(s,"&");
  for (var k in params) {
   var kv = String.split(params[k],"=");
   res[kv["0"]] = kv["1"];
  };
  return res;}
}
```

(The `encode` function is dual.) Our function illustrates support for arrays provided by our compiler,.Importantly, this code may be used to parse untrusted messages; our wrapper for `un` to `string` is straightforward—if the argument is already a string, it is just the identity. Hence, one can write efficient TS* that calls `decode` to parse messages received from the adversary; this coding style is idiomatic in JavaScript, while the checks performed by our type system and runtime prevent many pitfalls.

Another TS*sample illustrates the usage of `Rpc` and our typed JSON library (generalizing `QueryString`) to program a higher-level, statically typed API. It shows, for instance, how to program a client-side proxy for the "/me" method of the Facebook API, which retrieves the user profile; this TS* function has the return type:

```
type profile =
  {id: string; email: string; age_range: {min:number}; ... }
```

Cross-Domain Messaging The Facebook API is meant to run on any website and protects itself from a malicious host by using iframes. For example, if the website calls `FB.login`, the API loads an iframe from `facebook.com` that retrieves the current user's access token and then only sends it to the host website (via `postMessage`) if the host is in a list of authorized origins.

Bhargavan et al. (2013) report attacks on a prior version of this authorization code that were due to typing errors (and have now been patched). We re-implement this code in TS* and show how programmers can rely on typing to avoid such attacks.

The `checkOrigins` function below is given the current host origin and verifies it against an array of authorized origins. The `proxyMessage` function uses this check to guard the release of the token to the parent (host) website, using a safe copy of the primitive `postMessage` function.

```
function checkOrigins (given:string,expected:array string):bool{
 for (var k in expected) {
  if(given === expected[k]) return true;}
 return false;}
function proxyMessage(h:string,tok,origins:array string) {
 if(checkOrigins(h,origins)) postMessage('parent',tok,h);}
```

In a previous version of the Facebook API, `proxyMessage` was accidentally called with an `origins` parameter of type `string`, rather than `array string`. This innocuous type error leads to an attack, because the untyped version of the code succeeds with both strings and arrays, but with different results. To see the core problem, consider a call to `checkOrigins` where `given = "h"` and `expected = "http://W.com"`. The `for` loop iterates over each character of `expected`, and hence succeeds, when it should not. In TS*, iteration is well-typed only for arrays and, unlike JavaScript, this enumerates only the "own" properties of the array. Thus, in our code, this error is caught statically (if the incorrect call to `proxyMessage` is local) or dynamically (if the call is from another iframe); the check fails in both cases and the token is not leaked.

7. Conclusions and prospects

This paper aims to broaden the scope of gradual typing: not only it is useful for migrating dynamically type-safe code to more structured statically typed code, it is also useful for moving from unsafe code, vulnerable to security attacks, to a robust mixture of dynamically and statically type-safe code.

Within the context of JavaScript, we have presented TS*, a language with a gradual type system, a compiler, and runtime support that provides several useful safety and confinement properties. Our preliminary experience suggests that TS* is effective in protecting security-critical scripts from attacks—without safety and confinement, such properties are difficult to obtain for JavaScript, and indeed security for such scripts has previously been thought unobtainable in the presence of cross-site scripts.

Even excluding the adversary, TS* develops a new point in the design space of gradual typing, using an approach based on runtime type information. This has several useful characteristics, including a simple and uniform failure semantics, and its applicability to a language with extensible objects and object identity.

In the future, we plan to develop TS* along several dimensions. On the practical side, we expect to integrate our ideas in an experimental branch of the open source TypeScript compiler, targeting the construction of larger secure libraries. On the theoretical side, we plan to explore the formal certification of our compiler and runtime. We also hope to develop our preliminary ideas on new notions of blame to explain runtime failures in TS*.

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