Reasonably Programmable Literal Notation

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General-purpose programming languages typically define literal notation for only a small number of common data structures, e.g. lists. This is unsatisfying because there are many other data structures for which literal notation might be useful, e.g. finite maps, regular expressions, HTML data, SQL queries, syntax trees and chemical structures. There may also be different implementations of each of these data structures, perhaps with different performance characteristics, that could all benefit from common literal notation. This paper introduces *typed literal macros (TLMs)*, which allow library providers to define new literal notation of nearly arbitrary design at any specified type or parameterized family of types. Compared to existing approaches, TLMs are uniquely *reasonable*: TLM providers can reason modularly about syntactic ambiguity, and TLM clients can reason abstractly, i.e. without examining the underlying expansion, about types and binding. The system only needs to convey to clients, via secondary notation, the inferred *segmentation* of each literal, which gives the locations and types of spliced subterms. This paper establishes these abstract reasoning principles formally with a calculus of typed expressions, pattern matching and ML-style modules. This calculus is the first detailed type-theoretic account of a hygienic macro system, of any design, for a language with these essential features of ML. We are integrating TLMs into Reason, an emerging alternative front-end for OCaml.

1 Introduction

When designing the surface syntax of a general-purpose programming language, it is common practice to define shorthand *literal notation*, i.e. notation that decreases the syntactic cost of constructing and pattern matching over values of a particular data structure or parameterized family of data structures. For example, many languages in the ML family support list literals like `[x1, x2, x3]` in both expression and pattern position [25, 48]. While lists are common across problem domains, other literal notation is more specialized. For example, Ur/Web extends the surface syntax of Ur (an ML-like language [9]) with expression and pattern literals for encodings of XML and HTML data [10]. The example in Fig. 1 shows two HTML literals, one that "splices in" a string expression delimited by `{[` and `]}` and the other an HTML expression delimited by `{` and `}`.

```
1 fun heading first_name = <xml><h1>Hello, {[first_name]}!</h1></xml>
2 val body = <xml><body>{heading "World"} ...</body></xml>
```

Fig. 1. HTML literals with support for splicing at two different types are built primitively into Ur/Web [10].

This design practice, where the language designer privileges certain library constructs with built-in literal notation, is *ad hoc* in that it is easy to come up with other examples of data structures for which mathematicians, scientists or programmers have invented specialized notation [8, 32, 53]. For example, 1) clients of a "collections" library might want not just list literals, but also literal notation for matrices, finite sets, maps and so on; 2) clients of a "web programming" library might want CSS literals (which Ur/Web lacks); 3) a compiler author might want "quotation" literals for the terms of the object language and various intermediate languages of interest; and 4) clients of a "chemistry" library might want chemical structure literals based on the SMILES standard [4].

Although requests for specialized literal notation are easy to dismiss as superficial, the reality is that literal notation, or the absence thereof, can have a substantial influence on software quality. For example, Bravenboer et al. [6] finds that literal notation for structured encodings of queries, like the SQL-based query literals now found in many languages [46], reduce the temptation to use string encodings of queries and therefore reduce the risk of catastrophic string injection attacks [55]. More generally, evidence suggests that programmers frequently resort to "stringly-typed programming", i.e. they choose strings instead of composite data structures, largely for reasons of notational

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EXISTING GRAMMAR EXTENSION SYSTEMS

```
1 EXTEND /* loaded by, e.g., camlp4 */
2   expr:
3     | "`(" q = kquery ")`" -> q
4   kquery:
5     | /* ...K query grammar... */
6 /* ...more extensions defined... */
7 let w = compute_w();
8 let x = compute_x(w);
9 let y = `((!R)@&{&/x!/:2_!x}'!R)`;
```

(a) It is difficult to reason abstractly about programs that use unfamiliar grammar extensions (see the six reasoning criteria in the text).

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```
1 notation $kq at KQuery.t
2   lexer  KQueryLexer
3   parser KQueryParser.start
4   expansions require
5     module KQuery as KQuery;
6 /* ...more notations defined... */
7 let w = compute_w();
8 let x = compute_x(w);
9 let y = $kq `((!R)@&{&/x!/:2_!x}'!R)`;
```

(b) TLMs make examples like these more reasonable by leaving the base grammar fixed and making the type and binding structure and the segmentation explicit.

Fig. 2. Two of the possible ways to introduce literal notation for encodings of K queries

convenience. In particular, Omar et al. [53] sampled strings from open source projects and found that at least 15% of them could be parsed by some readily apparent type-specific grammar, e.g. for URLs, paths, regular expressions and many others. Literal notation, with support for splicing, would decrease the syntactic cost of composite encodings, which are more amenable to programmatic manipulation and compositional reasoning.

Of course, it would not scale to ask general-purpose language designers to build in support for all known notations *a priori*. Instead, there has been persistent interest in mechanisms that allow library providers to define new literal notation on their own. For example, direct grammar extension systems like Camlp4 [37] and Sugar* [17, 18], term rewriting systems like Template Haskell [43, 62], the system of *type-specific languages* (TSLs) described by Omar et al. [53], and other systems that we will discuss below can all be used to define new literal notation (and, in some cases, other forms of new notation, such as new infix expression forms, control flow operations or type declaration forms, which we leave beyond the scope of this paper).

Problem The problem that specifically motivates this paper is that these existing systems make it difficult or impossible to reason abstractly about such fundamental issues as types and variable binding when presented with a program using user-defined literal forms. Instead, programmers and editor services can only reason transparently, i.e. by inspecting the underlying expansion or the implementation details of the collection of extensions responsible for producing the expansion.

Consider, for instance, the perspective of a programmer trying to comprehend the program text in Fig. 2a, which is written in a popular dialect of OCaml's surface syntax called Reason [66] that has, hypothetically, been extended with some number of new literal forms by a grammar extension system — Lines 1-5 show the grammar extension syntax from Camlp4 [37], but Sugar*/SugarHaskell is similar [17–19]. Line 9 uses one of these active syntax extensions to construct an encoding of a query in the niche database query language K, using its intentionally terse notation [70]. The problem is that a programmer looking at the program as presented, and unfamiliar with (i.e. holding abstract) the details elided on Lines 5-6, cannot easily answer questions like the following:

up-to-date?

(1) **Responsibility**: Which syntax extension determined the expansion of the literal on Line 9? Might activating a new extension generate a conflicting expansion for the same literal?
(2) **Expansion Typing**: What type does y have?
(3) **Context Dependence**: Which bindings does the expansion of Line 9 invisibly depend on? If I shadow or remove a module or other binding, could that break or change the meaning of Line 9 because its expansion depends invisibly on the original binding?

(4) **Segmentation**: Are the characters `x`, `R` and `2` on Line 9 parsed as spliced terms, meaning that they appear directly in the underlying expansion, or are they parsed in some other way peculiar to this literal notation, e.g. as operators in the K query language?
(5) **Segment Typing**: What type is each spliced term expected to have? How can I infer a type for a variable that appears in a spliced term without looking at where it ends up in the expansion?
(6) **Capture**: If `x` is in fact a spliced term, does it refer to the binding of `x` on Line 5, or might it capture an invisible binding of the same identifier in the expansion of Line 9?

Forcing the programmer to reason transparently to answer basic questions like these defeats the ultimate purpose of syntactic sugar: decreasing cognitive cost [23]. Analagous problems do not arise when programming without syntax extensions in languages like ML — programmers can reason lexically about where variables and other symbols are bound, and types mediate abstraction over function and module implementations [59]. Ideally, the programmer would be able to abstract in some analagous manner over the implementation of an unfamiliar notation.

Given these issues, we concluded that direct grammar extension systems were not ideally suited for integration into the Reason platform, which seeks to develop a clear and modern surface syntax for the OCaml programming language. We also evaluated various approaches that are based not on direct grammar extension but on term rewriting over a fixed grammar. We give a full account of this evaluation in Sec. 9, but briefly, we found that:

- Unhygienic approaches like OCaml's preprocessor extension point (PPX) rewriters [37] and Template Haskell [43, 62] allow us to define new literal notation with support for splicing by repurposing existing string literal forms. They also partially or completely solve the problem of reasoning about **Responsibility** by using explicit annotations on terms that are being rewritten. However, they do not satisfy the remaining five reasoning criteria.
- Hygienic staging macro systems, like MetaOCaml and related systems [21, 34, 61] do not give the expander access to the syntax trees of arguments, i.e. the rewriting is parametric in the arguments, so they are not suitable for the problem of defining new literal notation.
- Hygienic term rewriting macro systems, like those in various Lisp-family languages [3, 11, 16, 20, 29, 30, 35, 45], e.g. Racket [20], as well as Scala [7], do not allow us to repurpose string literal forms to define composite literal forms because the hygiene discipline does not allow us to splice terms out of literal bodies via parsing (the system cannot identify these terms as sub-terms of the macro arguments [29, 30]).
- *Type-specific languages* (TSLs) [53] come closer to our goals in that they explicitly support splicing terms out of literal bodies, but the mechanism falls short with regard to the six reasoning principles just discussed in ways that we will return to in Sec. 9. In any case, this approach was designed for simple nominally-typed languages and relies critically on a particular local type inference scheme. It is not suitable for a language with an ML-like semantics, meaning a language with support for structural types (like tuple and function types in ML), parameterized type families, pattern matching and non-local type inference.

Contributions This paper introduces *typed literal macros* (TLMs): the first system for defining new literal notation that (1) provides the ability to reason abstractly about all six of the topics just outlined; and (2) is semantically expressive enough for integration into Reason/OCaml and other full-scale statically typed functional languages. We evaluate these claims with (1) a series of increasingly ML-like calculi equipped with proofs of the claimed abstract reasoning principles; and (2) a number of non-trivial examples appearing throughout the paper that involve the advanced language features mentioned above. In describing these examples, we also demonstrate that literal parsing logic can be defined using standard, unmodified parser generators, so the burden on notation providers is comparable to that of existing systems despite these stronger reasoning principles.

Brief Overview For a brief overview of our proposed mechanism, consider Fig. 2b. On Lines 1-5, we define a TLM named `$kq` that provides K query literal notation. On Line 9, we apply this TLM to express the example from Figure 2a. We can reason abstractly about this program as follows.

(1) **Responsibility**: The lexer and parser specified by the applied TLM on Lines 2-3 are together exclusively responsible for lexing, parsing and expanding the body of the generalized literal form, i.e. the characters between `` `( `` and `)` ``. We will give more details on generalized literal forms and on constructing a TLM lexer and parser in the next section. For now, let us simply reiterate that our design goal is to provide a system where the programmer does not normally need to look up the definitions of `KQueryLexer` and `KQueryParser` to reason about types and binding.
(2) **Expansion Typing**: The type annotation on the definition of `$kq` (Line 1) determines the type that the expansion must have, here `KQuery.t`.
(3) **Context Dependence**: Lines 4-5 specify that expansions generated by `$kq` are allowed to use the module `KQuery`, and no others. The system ensures that this dependency is bound as specified even if the variable `KQuery` has been shadowed at the application site. This completely relieves clients from needing to consider expansion-internal dependencies when naming variables.
(4) **Segmentation**: The intermediate output that the TLM generates is structured so that the system can infer from it an accurate *segmentation* of the literal body that distinguishes spliced terms, i.e. those that appear in the expansion, from segments parsed in some other way. The segmentation is all that needs to be communicated to language services downstream of the expander, e.g. editors and pretty printers, which can pass it on to the programmer using secondary notation, e.g. colors in this document. So by examining Line 9, the programmer knows that the two instances of `x` are spliced expressions (because they are in black), whereas the `R`'s must be parsed in some other way, e.g. as operators of the K language (because they are in green). This also implies that errors in spliced terms can always be reported in terms of their original location.
(5) **Segment Typing**: Each spliced segment in the inferred segmentation also has a type annotation. This, together with the context independence condition, ensures that type inference at the TLM application site can be performed (by editor services or in the programmer's mind) abstractly, i.e. by reference only to the type annotations on the spliced segments, not the full expansion.
(6) **Capture**: Splicing is guaranteed to be capture-avoiding, so the spliced expression `x` must refer to the binding of `x` on Line 2. It cannot have captured a coincidental binding of `x` in the expansion.

Paper Outline Sec. 2 details expression TLMs in Reason with a more detailed case study. Sec. 3 then introduces pattern TLMs and describes the special reasoning conditions in pattern position. Sec. 4 introduces the more general parametric TLMs, which allow us to define literal notation at a type- or module-parameterized family of types. This section also reveals that the **`expansions require`** clause can be understood as parameterization followed immediately by partial parameter application. Having introduced the basic machinery, we continue in Sec. 5 with examples that further demonstrate the expressive power of this approach and various parser implementation strategies. Notably, this section gives our take on the example from Fig. 1 of Ur/Web-style HTML literals as well as a TLM that implements quasiquotation for Reason language terms, which is useful for implementing other TLMs. Sec. 6 defines a type-theoretic calculus of simple expression and pattern TLMs and formally establishes the reasoning principles implied above in their simplest form. Sec. 7 adds type functions and an ML-style module system to this calculus, and gives a more general variant of the reasoning principles theorem. Sec. 8 provides a brief overview of how we are implementing TLMs without modifying OCaml's type system. Sec. 9 compares TLMs to related work, guided by the rubric of reasoning principles just discussed. Sec. 10 concludes with a summary of the contributions of this paper and a discussion of limitations and future work. Certain technical details, proofs and extensions to the presented calculi are available in the supplement.

```
module Regex = {
  type t = AnyChar
         | Str(string)
         | Seq(t, t)
         | Or(t, t)
         | Star(t);
};
```

(a) The Regex module, which defines the recursive datatype `Regex.t`.

```
module RegexNotation {
  notation $regex at Regex.t
    lexer   RegexLexer
    parser  RegexParser.start
    expansions require
      module Regex as Regex;
};
```

(b) The definition of the `$regex` TLM. Figure ?? defines `RegexLexer` and `RegexParser`.

```
1 open RegexNotation;
2 module DNA = { let any_base = $regex `(A|T|G|C)`; };
3 let bisA = $regex `(GC$(DNA.any_base)GC)`;
4 let restriction_template(gene) =
5   $regex `($(bisA)$(DNA.any_base)*$$(gene)$(DNA.any_base)*$(bisA))`;
```

(c) Examples of the `$regex` TLM being applied in a bioinformatics application.

Fig. 3. Case Study: POSIX-style regex literal notation, with support for string and regex splicing.

2 Expression TLMs in Reason

Consider the recursive datatype `Regex.t` defined by Fig. 3a, which encodes regular expressions (regexes) into Reason [67]. Regexes are common in, for example, bioinformatics, where they are used to express patterns in DNA sequences. We might construct a regex that matches the strings `"A"`, `"T"`, `"G"` or `"C"`, which represent the four bases in DNA, as follows:

```
let any_base = Regex.(Or(Str "A", Or(Str "T", Or(Str "G", Or(Str "C")))))
```

In Reason, the notation `Regex.(`e`)` locally opens the module `Regex` within the expression e, so we did not need to qualify each application of `Regex.Or` and `Regex.Str` above. Even with the aid of this shorthand, however, constructing regexes in this way is syntactically costly. Instead, we would like to have the option to use the common POSIX-style notation [1] when constructing values of type `Regex.t`, including values constructed compositionally from other regexes and strings. We solve this problem in Fig. 3b by defining a TLM named `$regex` (pronounced "lit regex") that supports POSIX-style regex notation extended with splice forms for regexes and strings. Fig. 3c shows three examples of `$regex` being applied.

2.1 Client Perspective

Let us start from the perspective of a client programmer examining Fig. 3 but holding the underlying expansion, as well as the details of the lexer and parser, `RegexLexer` and `RegexParser`, abstract. We will return to describe the lexer and parser from the provider's perspective in Sec. 2.2.

Line 2 of Fig. 3c applies `$regex` to construct the regex `DNA.any_base` that was described above, this time using the more clear and familiar POSIX regex notation. Line 3 applies `$regex` again, using its regex splice form to compositionally construct a regex matching DNA sequences recognized by the BisA restriction enzyme, where the middle base can be any base. Finally, Lines 4-5 of Fig. 3c define a function, `restriction_template`, that constructs a more complex regex from these first two regexes and a given gene sequence represented as a string.

Let us consider the second of these three TLM applications more closely:

```
$regex `(GC$(DNA.any_base)GC)`
```

According to the context-free grammar of (this paper's branch of) Reason, this form is simply a leaf of the unexpanded syntax tree, like a string literal would be. TLM names are prefixed by `$` to distinguish them from variables. We call the TLM argument, `` `(GC$(DNA.any_base)GC)` ``, a *generalized literal form*, following the terminology introduced in the prior work on type-specific languages [53].

The only lexical constraint on the *body* of a generalized literal form, i.e. the characters between `` `( `` and `` )` ``, is that any nested occurrences of `` `( `` must be balanced by `` )` ``, much like nested comments in Reason/OCaml. Generalized literal forms therefore lexically subsume other literal forms. The nesting constraint is to allow TLM applications to appear inside spliced expressions (discussed below). Prior work by Omar et al. [53] specified several other choices of outer delimitation, including layout-sensitive delimitation, but for the purposes of this work, we will use the outer delimiters `` `( `` and `` )` `` exclusively. Choices about how spliced expressions are recognized inside the literal body are entirely at the discretion of each TLM. In this case, notice that the TLM provider has chosen to use `$(`e`)` for regex splicing and `$$(`e`)` for string splicing, where e is an unexpanded Reason expression of arbitrary form. We return to the topic of splicing below.

Responsibility Responsibility for lexing, parsing and expanding each literal body is delegated uniquely to the lexer and parser specified in the definition of the applied TLM by the clauses **`lexer`** and **`parser`**, respectively. For now, let us purposefully continue to hold these details abstract.

TLM definitions follow the scoping structure of Reason, i.e. they can be defined in modules alongside other definitions. When a TLM definition appears inside a module, it must also appear in the signature with the same specification, up to the usual notions of type and module path equivalence in the type annotation and module dependencies, which are discussed below. This is much like the situation with datatype definitions in ML. In Sec. 8, we will discuss how we use an encoding trick involving nested modules and temporary exceptions internally in the implementation to avoid having to primitively add TLM definitions to OCaml's type and module system.

By convention, we define TLMs in a module suffixed with `Notation` so that client programmers can **`open`** just the TLM definition without bringing other definitions into scope, as shown on Line 1 of Fig. 3c. We could also have fully qualified each TLM application, `RegexNotation.$regex`, or abbreviated the TLM definition:

```
let $regex = RegexNotation.$regex;
```

What is fundamental about this design is that there is a well-defined and familiar protocol for finding the definition of the TLM uniquely responsible for each generalized literal form in a program, following the usual scoping rules of the language. An editor service or documentation tool can use this protocol to integrate TLM definition lookup into a "go to definition" command.

Expansion Typing Having found the definition of the `$regex` TLM, the client can immediately determine the type of the expansion being generated at each application site because it is specified explicitly by the clause **`at`** `Regex.t` on Line 2 of Fig. 3b. The expansion type of a TLM is analagous to the return type of a function. The identity of `Regex.t` is determined relative to the TLM definition site, not at each application site, so the module `Regex` need not be in scope at the application site, or it can have been shadowed by a different module.

Context Dependence The TLM mechanism enforces a strong context independence condition on generated expansions by requiring that the TLM definition explicitly specify all modules that the expansions it generates might internally depend on. In this case, Lines 5-6 of Fig. 3b specify that generated expansions might use the module `Regex`, again as it is bound at the TLM definition site, using the module variable `Regex` internally. In this case, the expansion-internal module variable happens to coincide with the module path, but in general, the dependency can be an arbitrary module path. For example, consider the following dependency clause:

```
expansions require
  module Regex as Regex
  module Core.List as List
```

The module variable `List` will be bound internally to `Core.List` in all expansions generated by this TLM, including expansions generated at application sites where `List` or `Core.List` are not

bound, or bound to different modules. All other bindings, whether at the TLM definition site or the TLM application site, are not internally available to the expansion. This allows both client programmers and the provider of the package where the TLM is defined to freely rename and shadow variables in the usual way, i.e. without needing to inspect each expansion or each TLM definition's parser to determine whether the syntax trees that it generates require certain variables be bound in particular ways at each application site (as is common with, e.g., syntax extensions based on `camlp4` or Sugar*, or when using unhygienic term rewriting approaches like OCaml's PPX system or Template Haskell, which we discuss further in Sec. 9).

Enforcing this strong context independence condition is technically quite subtle because TLM parsers need to be able to perform splicing, i.e. they need to be able to parse terms out of the literal body for placement in the expansion. Naïvely checking that only the explicitly named dependencies are free in the expansion would inappropriately enforce context independence on application site spliced expressions, which should certainly not be prevented from referring to variables in scope at the application site. For example, consider the final expansion of our example from Line 2 of Fig. 3c:

```
Regex.Seq(Regex.Str "GC", Regex.Seq(DNA.any_base, Regex.Str "GC"))
```

In this term, both `Regex` and `DNA` are free module variables. There is nothing to distinguish references to `DNA` that arose from a spliced sub-expression parsed out of the literal body from those that would indicate that the context independence condition has been violated.

As we will discuss more extensively in Sec. 9, this issue is why hygienic term-rewriting macro systems, like those available in various Lisp-family languages [45] and in Scala [7], cannot be used to repurpose string literals for composite literal notation at other types. These systems would find that the appearance of `DNA.any_base` in the generated expansion violates their context independence condition, because `DNA.any_base` is not, from the perspective of the context-free syntax, an argument to the macro nor even a sub-term of an argument for which a tree path can be assigned [29, 30]. Instead, it arises as the result of performing a complex operation—parsing—on a sub-string of the string literal provided as the macro argument.

To address this problem, TLM parsers are not tasked with generating the final expansion directly. Instead, the parser generates a *proto-expansion* that refers to spliced terms indirectly by their location relative to the start of the provided literal body. For example, the proto-expansion generated by `$regex` for the example above can be pretty printed as follows:

```
Regex.Seq(Regex.Str "GC", Regex.Seq(spliced<4; 15; Regex.t>, Regex.Str "GC"))
```

Here, `spliced<4; 15; Regex.t>` is a reference to the spliced expression `DNA.any_base` by its location relative to the start of the literal body being expanded—characters 4 through 15 (zero-indexed and inclusive) of the literal body `GC$(DNA.any_base)GC` are `DNA.any_base`. We say more about the type annotation on the splice reference when we discuss **Segment Typing** below. The context independence condition can be enforced straightforwardly on the proto-expansion – the only free variable in the proto-expansion is `Regex`, which is an explicitly listed dependency, so all is well. Had a reference to `DNA` appeared in the proto-expansion, then it would be clear that the corresponding final expansion could only be valid in certain contexts (where `DNA` is bound) and not others, and therefore this would be a violation of the context independence condition.

why not spliced[some-where

Segmentation The finite set of splice references in the proto-expansion generated for a literal body is called the *segmentation* of that literal body. The segmentation of the example above is the finite set containing one element, `spliced<4; 15; Regex.t>`. For the more complex example from Line 4 of Fig. 3c, the segmentation contains multiple spliced segments:

```
{ spliced<2; 5; Regex.t>, spliced<9; 20; Regex.t>; spliced<26; 29; string>;
  spliced<33; 44; Regex.t>; spliced<49; 52; Regex.t> }
```

The system checks that the segmentation does in fact segment the literal body, i.e. that the segments are in-bounds, of positive extent and non-overlapping. Adjacent spliced segments must also be separated by at least one character. The spliced segment locations can therefore be communicated unambiguously to the programmer by tools downstream of the expander, e.g. program editors and pretty printers, using secondary notation. In this paper, non-spliced segments are shown in color and spliced segments start in black.

When TLM applications are nested, a distinct color can be used at each depth. For example, in Sec. 5, we will describe a TLM `$html` for HTML notation in the style of Ur/Web (see Fig. 1). We can also define a TLM, `$smiles`, not shown, that allows for the use of the standard SMILES notation for chemical structures [4], extended with splicing notation. Assuming that the expansion type of `$smiles` is `Smiles.t` and that `Smiles.to_svg : Smiles.t -> Html.t`, we can nicely transform encodings of chemical structures into a vector graphic embedded into an HTML page as follows:

```
$html `(
  <div>Chemical structure of sucrose: {
    Smiles.to_svg
      $smiles `({mono_glucose}-O-{mono_fructose})`
  }</div>
)`
```

Another benefit of explicitly tracking the locations of spliced segments is that errors that originate in spliced terms can be reported in terms of their original source location [68].

Segment Typing Each splice reference in the segmentation specifies not just the location of a spliced expression but also its expected type. The identity of this type is resolved in a context-independent manner, assuming only the dependencies explicitly specified by the TLM.

By associating a type with each spliced segment, we guarantee that type inference can be performed abstractly, meaning that only the segment types, together with the expansion types specified by the applied TLMs, are necessary to infer types for variables appearing in a client-side function. For example, consider the function `restriction_template` on Lines 3-4 of Fig. 3c. The return type of this function can be inferred to be `Regex.t` from the type annotation on the `$regex` TLM, as previously discussed. The type of the argument, `gene`, can be inferred to be **`string`** because the segmentation specifies the type **`string`** for the spliced segment where it appears (cf. the segmentation shown under **Segmentation** above). The context independence condition implies that `gene` cannot appear elsewhere in the expansion, and so no further typing constraints could possibly be collected from examining the portions of the (proto-)expansion being held abstract.

Segment types can be communicated directly to the programmer upon request by an editor service. For Reason, we are extending the Merlin editor service [2], which is used by various Reason editor extensions (e.g. for Emacs, Vim, and so on), with a new editor command for reporting the expected type of the innermost spliced segment containing the cursor. Note that because the type is specified explictly in the output of the parser, this information can be reported even when there is a parse error or type error in a spliced expression.

Segment types are somewhat analagous to the argument types of a function. The difference is that the argument signature of a function is the same every time the function is applied, i.e. it is associated with the function itself, whereas the inferred segmentation can differ at each TLM application site. This is, of course, what gives TLMs their notational flexibility.

Capture In discussing the question of inferring a type for `gene` above, we neglected to consider one critical question: are we sure that the variable `gene` in the third spliced segment on Line 4 of Fig. 3c is, in fact, a reference to the argument `gene` of the `restriction_template` function? After all, if we hold the expansion abstract then it may well be that the third spliced segment appears under

(i.e. captures) a different binding of the identifier `gene`. For more common identifiers, e.g. `x` and `tmp`, inadvertent capture is perhaps inevitable.

For example, consider the following application site:

```
let tmp = /* ... application site temporary ... */;
$html `(<h1>{f(tmp)}</h1>)`;
```

Now consider the scenario where the proto-expansion generated by `$html` has the following form:

```
let tmp = /* ... expansion-internal temporary ... */;
Html.H1Element(tmp, spliced<5; 10; Html.t>);
```

If the final expansion was produced naïvely, by syntactically replacing the spliced segment reference by the final expansion recursively determined for the corresponding spliced expression, then the variable `tmp` in the spliced expression would capture the binding of `tmp` in the proto-expansion. The result when the types of the two bindings differed would be a type error that exposes the internal details of the expansion. If the types of the two bindings of `tmp` coincided, then there could be subtle changes in run-time behavior that would be impossible to diagnose without looking at the expansion itself. Capture of invisible bindings is, due to these hazards, clearly unreasonable.

To address this problem in our design, splicing is guaranteed to be capture-avoiding. The final expansion is generated by recursively expanding each spliced expression and then inserting it into the final expansion via capture-avoiding substitution. In other words, the system automatically alpha-varies the bindings in the proto-expansion to ensure capture avoidance. There is no need for TLM providers to manually deploy a mechanism that generates fresh variables (as in, e.g., Racket's *reader macro* system [20], which is further discussed in Sec. 9). For example, the final expansion of the example above would be alpha-equivalent to the following:

```
let tmp = /* ... application site temporary ... */;
let tmp' = /* ... expansion-internal temporary (automatically renamed) ... */;
Html.H1Element(tmp', f(tmp));
```

Although this capture avoidance discipline implies that TLMs cannot intentinoally expose bindings internal to the expansion directly to spliced expressions, this does not imply that values cannot flow from the expansion into a spliced expression. It simply means that when this is intended, the segment type must be a function type. Reason has concise lambda notation, `(x) => e`, which is useful in many of these use cases. For example, the capture avoidance discipline means that we cannot define numeric list comprehension notation[1] like the following:

```
$listcomp `(x + 1 | x in lo to hi)` (* NO! cannot reason abstractly *)
```

However, the following is permitted, because `x` is bound by the spliced lambda expression:

```
$listcomp `((x) => x + 1 | lo to hi)`  (* OK! *)
```

Our contention is that small syntactic costs like these are more than justified by the peace of mind of knowing that unfamiliar literal notation is absolutely prevented from obscuring the type and binding structure of the language. We say more about the future prospect of a mechanism designed specifically for specialized binding forms, e.g. Haskell-style **do** notation [33], in Sec. 10.

do this

Summary Let us briefly reiterate the key ideas of this section. The focus was on providing simple protocols to client programmers when they have questions about the syntactic structure, type structure or binding structure of a program. At no point in answering these sorts of questions is the client made to look at the generated expansion itself, nor inspect the parser implementation. Instead, the programmer need only be given knowledge of the expansion type from the TLM definition and the segmentation inferred by the expander at each application site, which carries a small volume of information that can be communicated straightforwardly by standard editor

[1] A polymorphic list comprehension TLM requires type parameters, introduced in Sec. 4.

```
type body = string;
type segment = {startIdx: int, endIdx: int};
type parse_result('a)
  = ParseError {msg: string, loc: segment}
  | Success('a);
type proto_typ = Arrow(proto_typ, proto_typ)
               | StringTy
               | /* ... */
               | SplicedT(segment);
type proto_expr = Tuple(list(proto_expr))
                | /* ... */
                | SplicedE(segment, proto_typ);
```

Fig. 4. Definitions of various types available ambiently to TLM definitions.

services. Certain questions related to binding structure simply do not need to be asked due to the strict context independence and capture avoidance discipline of the system, enabled by the explicit tracking of spliced segments by location. Despite these semantic constraints, the system is able to express a number of interesting examples with few compromises because there is only one simple lexical constraint on literal bodies.

————————- stuff below is notes / not yet revised

2.2 Provider Perspective

If the parse function determines that the literal body is not well-formed according to the syntax that it implements, it must return `ParseError {msg=`e_{msg}`, loc=`e_{loc}`}` where e_{msg} is a custom error message and e_{loc} is a value of type `segment`, defined in Figure 4, that designates a segment of the literal body as the origin of the error [68].

If instead parsing succeeds, the parse function returns `Success` e_{proto}, where e_{proto} is called the *encoding of the proto-expansion*. For expression TLMs, the proto-expansion is a *proto-expression* and it is encoded as a value of the recursive datatype `proto_expr` that is outlined in Figure 4. Most of the constructors of `proto_expr` are individually uninteresting – they encode OCaml's various expression forms. Expressions can mention types, so we also need the type `proto_typ` also outlined in Figure 4. It is only the `SplicedE` and `SplicedT` constructors that are novel. These are discussed next.

The fact that the context-free grammar of the base language is fixed ensures that notation providers can reason modularly about syntactic ambiguity in their own grammars.

2.3 Splicing

When the parse function determines that some segment of the literal body is a spliced expression, according to whatever syntactic criteria it deems suitable, it can indirectly refer to it in the encoding it produces using the `SplicedE` constructor of `proto_expr`, which takes a value of type `segment` that indicates the zero-indexed location of the spliced expression relative to the start of the provided literal body. The `SplicedE` constructor also requires a value of type `proto_typ`, which indicates the type that the spliced expression is expected to have. Types can be spliced out by using the `SplicedT` constructor of `proto_typ` analagously.

For example, consider again the two TLM applications in Figure 5c. In each case, the parse function of the `$html` TLM (Figure 5b, Lines 2-4) first sends the literal body through an off-the-shelf HTML parser, `parse_html`. It then passes the result to a function `html_to_ast`, not shown, which produces the corresponding expression encoding of type `proto_expr`. When this function

encounters an HTML text node containing matched `{[` and `]}`, then that segment is inserted as a spliced expression of type **`string`**, and similarly text nodes containing matched curly braces produce a spliced expression of type `html`. For instance, the proto-expansion generated for first TLM application in 5c, pretty-printed, is:

```
H1Element(Nil, Cons(TextNode "Hello, ", Cons(
  TextNode spliced<13; 22; string>, Nil)))
```

Here, **`spliced`**`<13; 22;` **`string`**`>` is a reference to the spliced string expression `first_name` by its location relative to the start of the literal body being expanded (the off-the-shelf HTML parser provides the necessary baseline location information for use by `html_to_ast`). It corresponds to the encoding `SplicedE({startIdx=13, endIdx=22}, StringTy)`. Requiring that TLMs refer to spliced expressions indirectly in this manner ensures that a TLM cannot "forge" spliced terms, i.e. claim that some sub-term of the expansion should be given the privileges of a spliced term, discussed in Sec. 2.4, when it does not in fact appear in the literal body.

The *segmentation* inferred from a proto-expansion is the finite set of references to spliced terms contained within. For example, the segmentation inferred from the proto-expression above contains only **`spliced`**`<13; 22;` **`string`**`>`. The system checks that all of the locations in the segmentation are 1) in bounds relative to the literal body; and 2) non-overlapping. This resolves the problem of **Segmentation** described in Sec. 1, i.e. every literal body in a well-formed program has a well-defined segmentation. The TSL mechanism did not maintain this reasoning principle [53]. A program editor or pretty-printer can communicate this segmentation information to the programmer, e.g. by coloring non-spliced segments green as is our convention in this document. In general, spliced expressions might themselves apply TLMs, in which case the convention is to use a distinct color for unspliced segments at each depth. For example, consider a TLM `$smiles` for chemical structures [4] with support for splicing using curly braces:

```
$html [|Chemical structure of sucrose: {
    $smiles [|{m_glucose}-O-{m_fructose}|]
    |> SMILES.to_svg }|]
```

A program editor or pretty-printer can communicate the type of each spliced expression, also specified abstractly by the segmentation, upon request (for Reason, via Merlin [2].)

2.4 Proto-Expansion Validation

Three important concerns described in Sec. 1 remain: those related to reasoning abstractly about the *hygiene properties*, i.e. **Capture** and **Context Dependence**, and **Typing**. Addressing these concerns is the purpose of the *proto-expansion validation* process, which occurs during the typing phase. Proto-expansion validation results in the *final expansion*, which is simply the proto-expansion with the references to spliced segments replaced with their own final expansions.

2.4.1 Capture. Proto-expansion validation ensures that spliced terms have access *only* to the bindings at the application site—spliced terms cannot capture bindings internal to the proto-expansion.

2.4.2 Context Dependence. The proto-expansion validation process also ensures that variables that appear in the proto-expansion do not refer to bindings that appear either at the TLM definition or the application site. In other words, expansions must be completely *context independent* – they can make no assumptions about the surrounding context whatsoever.

A minimal example of a "broken" TLM that does not generate context-independent proto-expansions is below:

```
syntax $broken at t by static {
```

```
    fun(_) => Success (Var "SSTRxESTR") };
```

The proto-expansion that this TLM generates (for any literal body) refers to a variable `x` that it does not itself bind. If proto-expansion validation permitted such a proto-expansion, it would be well-typed only under those application site typing contexts where `x` is bound. This "hidden assumption" makes reasoning about binding and renaming difficult.

Of course, this prohibition does not extend into the spliced terms in a proto-expansion – spliced terms appear at the application site, so they can justifiably refer to application site bindings. (like `first_name` in Fig. 5c.) Because proto-expansions refer to spliced terms indirectly, enforcing context independence is straightforward – we need only that the proto-expansion itself be closed.

Naïvely, this restriction, also present in the prior work on TSLs [53], is quite restrictive – expansions cannot access any library functions. At best, they can require the client to "pass in" required library functions via splicing at every application. In Sec. 4, we will introduce module parameters and partial parameter application to neatly resolve this problem.

2.4.3 Typing. Finally, proto-expansion validation maintains a reasonable typing discipline by (1) checking that the expansion is of the type specified by the TLM's type annotation; (2) checking that each spliced type is valid; (3) checking that the type annotation on each spliced expression is valid; and (4) checking each spliced expression against the specified type annotation. Context independence implies that ML-style type inference can be performed using only the segmentation (because the remainder of an expansion cannot mention the very variables whose types are being inferred). In the prior work on TSLs, spliced terms did not have type annotations

3 Pattern TLMs

Let us now briefly consider the topic of TLMs that generate *patterns*, rather than expressions. Pattern literals are the dual to expression literals in that expression literals support construction whereas pattern literals support deconstruction [50]. For example, we can pattern match on a value `x : html` by applying a pattern TLM `$html` as follows:

```
switch x {
| $html [|<h1>{cs}</h1>|] => /* cs:list(html) */
| _ -> None
};
```

Any list pattern, including one generated by another TLM application, can appear where `cs` appears in the example pattern above. For longer **`switch`** expressions, the shorthand **`switch`** `x using $html` applies `$html` to every rule where the outermost pattern is of generalized literal form.

Notice that we can use the same name, `$html`, for this pattern TLM as for the expression TLM defined in the previous section. It does not make sense to apply an expression TLM in pattern position (many expression-level constructs, e.g. lambdas, do not correspond even syntactically to patterns), and *vice versa*, so this is unambiguous.

Pattern TLM definitions differ from expression TLM definitions in two ways: (1) a *sort qualifier*, **`for patterns`**, distinguishes them from expression TLM definitions; (2) the return type of the parse function is `parse_result(proto_pat)`, rather than `parse_result(proto_expr)`. The type `proto_pat`, outlined below, classifies encodings of *proto-patterns*.

```
type proto_pat = /* no variable pattern form! */
               | Wild
               | /* ... */
               | SplicedP(segment, proto_typ);
```

The constructor `SplicedP` operates much like `SplicedE` to allow a proto-pattern to refer indirectly to spliced patterns.

To maintain the abstract binding discipline, variable patterns can appear only within spliced patterns. Enforcing this restriction is straightforward: we simply have not defined a variant of the `proto_pat` type that encodes variable patterns (wildcards are allowed.) This restriction ensures that only variables visible to the client in a spliced pattern are bound in the corresponding branch expression. This is analagous to the capture avoidance principle for expression TLMs.

Type annotations on references to spliced patterns could refer to type variables, so we also need to enforce context independence in the manner discussed in the previous section.

To maintain an abstract typing discipline, proto-pattern validation checks type annotations much as in Sec. 2.4.3.

4 Parametric TLMs

The simple TLMs in the previous sections operate only at one specified type, as did the TSLs in the prior work [53]. This is rather limiting. This section introduces *parametric TLMs*, which can operate over a type- and module-parameterized family of types. They also neatly solve the problem discussed in Sec. 2.4.2 of giving expansions access to helper functions.

Consider the following Reason/OCaml module type (a.k.a. signature), which specifies an abstract data type [25, 38] of string-keyed polymorphic dictionaries:

```
module type DICT = {
  type t('a);
  let empty : t('a);
  let extend : t('a) -> (string, 'a) -> t('a);
  /* ... */
};
```

We can define a TLM that is parametric over implementations of this signature, `D : DICT`, and over choices of the codomain type, `'a` as follows:

```
syntax $dict (D : DICT) (type 'a) at D.t('a)
by static { fun(b) => /* ... */ };
```

For example, given some module that implements this signature, `HashDict : DICT`, we can apply `$dict` as follows:

```
$dict HashDict int [|"key1"=>10; "key2"=>15|]
```

Notice that the segmentation immediately reveals which punctuation is particular to this TLM and where the spliced key and value expressions appear. Because the context-free syntax of unexpanded terms is never modified, it is possible to reason modularly about syntactic determinism, i.e. we can reason above that => does not appear in the follow set of unexpanded expressions [60], so there can never be an ambiguity about where a key expression ends.

The proto-expansion generated for the TLM application above, shown below, can refer to the parameters:

```
D.extend(D.extend D.empty (spliced<1;6;string>,
  spliced<11;12;'a>)) (spliced<15;20;string>,
  spliced<25;26;'a>)
```

Validation checks that the proto-expansion is truly parametric, i.e. it must be valid for all modules `D : DICT` and types `'a`. It is only after validation that we substitute the actual parameters, here `HashDict` for `D` and `int` for `'a`, into the final expansion. Only the type annotations on references to spliced terms are subject to early parameter substitution (because they classify application site terms.)

If we will use the `HashDict` implementation ubiquitously, we can abbreviate the partial application of `$dict` to `HashDict`, resulting in a TLM that is parametric over only the type `'a`:

```
type html_attrs = list(string, string)
type html = BodyElement(html_attrs, list(html))
          | H1Element(html_attrs, list(html))
          | TextNode(string)
          | /* ... */;
```

(a) The `html` type, which classifies encodings of HTML data.

```
syntax $html at html by static {
  fun (b : body) : parse_result(proto_expr) =>
    try (parse_html b |> html_to_ast |> Success
    | InvalidHTML msg loc => ParseError msg loc
};
```

(b) The `$html` TLM definition. Figure 4 defines TLM-related types.

```
let heading first_name = $html [|<h1>Hello, {[first_name]}!</h1>|]
let body = $html [|<body>{heading "World!"} ...</body>|]
```

(c) Two examples of the `$html` TLM being applied. Compare to Ur/Web's built in HTML literal notation from Figure 1.

Fig. 5. Case Study: HTML literals

```
syntax $dict' = $dict HashDict;
```

TLM abbreviations can themselves be parameterized to support partial application of parameters other than the last.

In Sec. 2.4.2 we discussed the problem of the strict context independence discipline being too restrictive, in that it would seem to restrict expansions from referring to useful helpers bound at the TLM definition site. Module parameters address this problem – the helper values, types and modules can be packaged into a module, passed in and partially applied to hide this detail from clients. Because this will be common in practice, we provide the following shorthand:

```
syntax $a at t using X~${}_1$~=M~${}_1$~,...,X~${}_n$~=M~${}_n$~ by ...
```

This explicit parameter passing discipline is reminiscent of work on explicit capture for distributed functions [47]. By not implicitly giving expansions access to all definition-site bindings, we need not examine parse functions to reason about, e.g., renaming. Consequently, encodings of proto-terms can be values of standard datatypes (e.g. `proto_expr`) with variables represented simply as, e.g., strings, and quasiquotation notation can be expressing using TLMs (see supplement). A central design goal of Reason is to leave the OCaml semantics unchanged, so building quasiquotation in primitively, to integrate the free variables in term encodings into the overall binding discipline [5], was in any case infeasible.

5 Additional Examples

5.1 HTML

We start in this section by describing the TLM mechanism in detail by way of a substantial case study: literal notation, like that built in to Ur/Web (see Figure 1) and also, presently, Reason, for expressions of the type `html` in Figure 5a.

5.2 Quotation

5.3 More Examples, Briefly

6 Simple TLMs, Formally

This section will present a calculus of simple expression and pattern TLMs called $\lambda^{\mathbf{S}}_{\mathbf{TLM}}$. By the end of this section, we will have a theorem that encodes the six reasoning principles that were outlined informally in the previous sections.

Because our focus is on these reasoning principles, and for reasons of space, we leave our full calculus of parametric TLMs to the supplement. The full calculus extends the simple calculus of this section with an ML module calculus based closely on the system defined by Harper [26], which

$$
\begin{array}{lrcl}
\text{UTyp} & \hat{\tau} & ::= & \hat{t} \mid \hat{\tau} \rightharpoonup \hat{\tau} \mid \forall \hat{t}.\hat{\tau} \mid \mu \hat{t}.\hat{\tau} \\
& & \mid & \langle \{i \hookrightarrow \hat{\tau}_i\}_{i \in L} \rangle \mid [\{i \hookrightarrow \hat{\tau}_i\}_{i \in L}] \\
\text{UExp} & \hat{e} & ::= & \hat{x} \mid \lambda \hat{x}{:}\hat{\tau}.\hat{e} \mid \hat{e}(\hat{e}) \mid \Lambda \hat{t}.\hat{e} \mid \hat{e}[\hat{\tau}] \mid \mathtt{fold}(\hat{e}) \\
& & \mid & \langle \{i \hookrightarrow \hat{e}_i\}_{i \in L} \rangle \mid \mathtt{inj}[\ell](\hat{e}) \mid \mathtt{match}\ \hat{e}\ \{\hat{r}_i\}_{1 \leq i \leq n} \\
& & \mid & \mathtt{syntax}\ \hat{a}\ \mathtt{at}\ \hat{\tau}\ \mathtt{for\ expressions\ by\ static}\ e\ \mathtt{in}\ \hat{e} \\
& & \mid & \mathtt{syntax}\ \hat{a}\ \mathtt{at}\ \hat{\tau}\ \mathtt{for\ patterns\ by\ static}\ e\ \mathtt{in}\ \hat{e} \\
& & \mid & \hat{a}\ [\![b]\!] \\
\text{URule} & \hat{r} & ::= & \hat{p} \Rightarrow \hat{e} \\
\text{UPat} & \hat{p} & ::= & \hat{x} \mid \_ \mid \mathtt{fold}(\hat{p}) \mid \langle \{i \hookrightarrow \hat{p}_i\}_{i \in L} \rangle \mid \mathtt{inj}[\ell](\hat{p}) \mid \hat{a}\ [\![b]\!]
\end{array}
$$

Fig. 6. Syntax of the $\lambda^{\mathbf{S}}_{\mathbf{TLM}}$ unexpanded language (UL). Metavariable $\hat{t}$ ranges over type identifiers, $\hat{x}$ over expression identifiers, ℓ over labels, L over finite sets of labels, $\hat{a}$ over TLM names and b over literal bodies. We write $\{i \hookrightarrow \hat{\tau}_i\}_{i \in L}$ for a finite mapping of each label i in L to some unexpanded type $\hat{\tau}_i$, and similarly for other sorts. We write $\{\hat{r}_i\}_{1 \leq i \leq n}$ for a finite sequence of n unexpanded rules.

in turn is based on early work by MacQueen [41, 42], subsequent work on the phase splitting interpretation of modules [27] and on using dependent singleton kinds to track type identity [12, 64], and finally on formal developments by Dreyer [15] and Lee et al. [36]. These additional mechanisms are necessary only to formalize the advanced features of Sec. 4. Proofs are in the supplement for both the simple and full calculus.

Both the simple and full calculus consist of an *unexpanded language*, or *UL*, defined by typed expansion to an *expanded language*, or *XL*. Figs. 6 and 7 summarize the syntax of the UL and the XL, respectively.

6.1 Expanded Language (XL)

The XL of $\lambda^{\mathbf{S}}_{\mathbf{TLM}}$ forms a standard pure functional language with partial function types, quantification over types, recursive types, labeled product types and labeled sum types and support for pattern matching. The reader is directed to *PFPL* [26] for a detailed introductory account of these constructs. We will only tersely summarize the statics and dynamics of the XL below because the particularities are not critical.

The *statics of the XL* is organized around the type formation judgement, $\Delta \vdash \tau$ type, the expression typing judgement, $\Delta\ \Gamma \vdash e : \tau$, and the pattern typing judgement, $\Delta \vdash p : \tau \dashv \Gamma$. In the latter, Γ is a typing context that tracks the typing hypotheses generated by p. These judgements are inductively defined in the supplemental material along with necessary auxiliary structures and standard lemmas.

The *evaluation semantics* of $\lambda^{\mathbf{S}}_{\mathbf{TLM}}$ is organized around the judgements e val, which says that e is a value, and $e \Downarrow e'$, which says that e evaluates to the value e'.

6.2 Syntax of the Unexpanded Language

Unexpanded types and expressions are simple inductive structures. Unlike expanded types and expressions, they are **not** abstract binding trees – we do **not** define the standard notions of renaming, alpha-equivalence or substitution for unexpanded terms. This is because unexpanded expressions remain "partially parsed" due to the presence of literal bodies, b, from which spliced terms might be extracted during typed expansion. In fact, unexpanded types and expressions do not involve variables at all, but rather *type identifiers*, $\hat{t}$, and *expression identifiers*, $\hat{x}$. Identifiers are given meaning by expansion to variables during typed expansion, as we will see. This distinction between identifiers and variables is technically crucial to our developments.

Most of the unexpanded forms in Figure 6 mirror the expanded forms. We refer to these as the *common forms*.

$$
\begin{array}{lrcl}
\text{Typ} & \tau & ::= & t \mid \mathtt{parr}(\tau;\tau) \mid \mathtt{all}(t.\tau) \mid \mathtt{rec}(t.\tau) \\
& & \mid & \mathtt{prod}[L](\{i \hookrightarrow \tau_i\}_{i\in L}) \mid \mathtt{sum}[L](\{i \hookrightarrow \tau_i\}_{i\in L}) \\
\text{Exp} & e & ::= & x \mid \mathtt{lam}\{\tau\}(x.e) \mid \mathtt{ap}(e;e) \mid \mathtt{tlam}(t.e) \mid \mathtt{tap}\{\tau\}(e) \\
& & \mid & \mathtt{fold}(e) \mid \mathtt{tpl}[L](\{i \hookrightarrow e_i\}_{i\in L}) \mid \mathtt{inj}[\ell](e) \\
& & \mid & \mathtt{match}[n](e;\{r_i\}_{1\le i\le n}) \\
\text{Rule} & r & ::= & \mathtt{rule}(p.e) \\
\text{Pat} & p & ::= & x \mid \mathtt{wildp} \mid \mathtt{foldp}(p) \mid \mathtt{tplp}[L](\{i \hookrightarrow p_i\}_{i\in L}) \\
& & \mid & \mathtt{injp}[\ell](p)
\end{array}
$$

Fig. 7. Syntax of the expanded language (XL). XL terms are *abstract binding trees* (ABTs) identified up to alpha-equivalence, so we follow the syntactic conventions of Harper [26]. Metavariable x and t ranges over variables.

$$
\dfrac{\dfrac{}{\hat\Delta \vdash \hat\tau \rightsquigarrow \tau \text{ type}} \qquad \dfrac{\dfrac{}{\hat\Delta\ \langle \hat x \rightsquigarrow x_2; x_1 : \tau, x_2 : \tau\rangle \vdash_{\hat\Psi;\hat\Phi} \hat x \rightsquigarrow x_2 : \tau}\ \textsc{ee-id}}{\hat\Delta\ \langle \hat x \rightsquigarrow x_1; x_1 : \tau\rangle \vdash_{\hat\Psi;\hat\Phi} \lambda\hat x{:}\hat\tau.\hat x \rightsquigarrow \mathtt{lam}\{\tau\}(x_2.x_2) : \mathtt{parr}(\tau;\tau)}\ \textsc{ee-lam}}{\hat\Delta\ \langle \emptyset;\emptyset\rangle \vdash_{\hat\Psi;\hat\Phi} \lambda\hat x{:}\hat\tau.\lambda\hat x{:}\hat\tau.\hat x \rightsquigarrow \mathtt{lam}\{\tau\}(x_1.\mathtt{lam}\{\tau\}(x_2.x_2)) : \mathtt{parr}(\tau;\mathtt{parr}(\tau;\tau))}\ \textsc{ee-lam}
$$

Fig. 8. An example expansion derviation demonstrating how identifiers and variables are separately tracked.

There is also a corresponding context-free textual syntax for the UL. Giving a complete definition of the context-free textual syntax as, e.g., a context-free grammar, is not critical to our purposes here. Instead, we only posit partial metafunctions parseUTyp(b), parseUExp(b) and parseUPat(b) that go from character sequences, b, to unexpanded types, expressions and patterns (the supplement states the full condition.)

6.3 Typed Expansion

Unexpanded terms are checked and expanded simultaneously according to the central *typed expansion judgements*:

$$
\begin{array}{ll}
\hat\Delta \vdash \hat\tau \rightsquigarrow \tau \text{ type} & \hat\tau \text{ has well-formed expansion } \tau \\
\hat\Delta\ \hat\Gamma \vdash_{\hat\Psi;\hat\Phi} \hat e \rightsquigarrow e : \tau & \hat e \text{ has expansion } e \text{ of type } \tau \\
\hat\Delta \vdash_{\hat\Phi} \hat p \rightsquigarrow p : \tau \dashv\!\!\dashv \hat\Gamma & \hat p \text{ has expansion } p \text{ matching } \tau
\end{array}
$$

The typed expansion rules that handle common forms mirror the corresponding typing rules. The *expression TLM context*, $\hat\Psi$, and the *pattern TLM context*, $\hat\Phi$, pass through these rules opaquely. For example, the rules for variables and lambdas are shown being applied in Fig. 8, discussed below.

The only subtlety related to common forms has to do with the relationship between identifiers, $\hat x$, in the UL and variables, x, in the XL. To understand this, we must first describe in detail how unexpanded contexts work. *Unexpanded typing contexts*, $\hat\Gamma$, are pairs of the form $\langle\mathcal{G};\Gamma\rangle$, where $\mathcal{G}$ maps each expression identifier $\hat x \in \mathrm{dom}(\mathcal{G})$ to the hypothesis $\hat x \rightsquigarrow x$, for some expression variable, x, called its expansion. The standard typing context, Γ, then tracks the type of x. We write $\mathcal{G} \uplus \hat x \rightsquigarrow x$ for the expression identifier expansion context that maps $\hat x$ to $\hat x \rightsquigarrow x$ and defers to $\mathcal{G}$ for all other expression identifiers (i.e. the previous mapping is **updated**.) Note the distinction between update and extension (which requires that the new identifier is not already in the domain.) We define $\hat\Gamma, \hat x \rightsquigarrow x : \tau$ when $\hat\Gamma = \langle\mathcal{G};\Gamma\rangle$ as an abbreviation of $\langle\mathcal{G} \uplus \hat x \rightsquigarrow x; \Gamma, x : \tau\rangle$.

To develop an intuition for why the update operation is necessary, it is instructive to inspect in Fig. 8 the derivation of the expansion of the unexpanded expression $\lambda\hat x{:}\hat\tau.\lambda\hat x{:}\hat\tau.\hat x$ to $\mathtt{lam}\{\tau\}(x_1.\mathtt{lam}\{\tau\}(x_2.x_2))$ assuming $\hat\Delta \vdash \hat\tau \rightsquigarrow \tau$ type. Notice that when Rule EE-LAM is applied, the type identifier expansion context is updated but the typing context is extended with a

$$\textsc{ee-def-setsm}\quad \frac{\hat{\Delta} \vdash \hat{\tau} \rightsquigarrow \tau \text{ type} \qquad \emptyset\,\emptyset \vdash e_{\text{parse}} : \texttt{parr(Body;ParseResultSE)} \qquad e_{\text{parse}} \Downarrow e'_{\text{parse}} \qquad \hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi}, \hat{a} \rightsquigarrow a \hookrightarrow \texttt{setlm}(\tau; e'_{\text{parse}}); \hat{\Phi}} \hat{e} \rightsquigarrow e : \tau'}{\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi}; \hat{\Phi}} \texttt{syntax}\ \hat{a}\ \texttt{at}\ \hat{\tau}\ \texttt{for expressions by static}\ e_{\text{parse}}\ \texttt{in}\ \hat{e} \rightsquigarrow e : \tau'}$$

$$\textsc{ee-ap-setsm}\quad \frac{b \downarrow_{\text{Body}} e_{\text{body}} \qquad e_{\text{parse}}(e_{\text{body}}) \Downarrow \texttt{inj[Succ} \qquad e_{\text{proto}} \uparrow_{\text{PrExpr}} \grave{e} \qquad \text{seg}(\grave{e}) \text{ segme} \qquad \emptyset\,\emptyset \vdash^{\hat{\Delta};\hat{\Gamma};\hat{\Psi};\hat{\Phi};b} \grave{e} \rightsquigarrow e : \tau}{\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi}', \hat{a} \rightsquigarrow a \hookrightarrow \texttt{setlm}(\tau; e_{\text{parse}}); \hat{\Phi}} \hat{a}\ [\![b]\!]}$$

Fig. 9. The typed expansion rules for expression TLM definition and application.

(necessarily fresh) variable, first x_1 then x_2. Without this mechanism, expansions for unexpanded terms with shadowing, like this minimal example, would not exist, because we cannot implicitly alpha-vary the unexpanded term to sidestep this problem in the usual manner.

6.4 TLM Definitions

Rule EE-DEF-SETSM in Fig. 9 governs simple expression TLM (seTLM) definitions. The first premise expands the unexpanded type annotation. The second premise checks that e_{parse} is a closed expanded function of the given function type. (In the supplement, we add the machinery necessary for parse functions that are neither closed nor yet expanded.)

The type abbreviated `Body` classifies encodings of literal bodies, b. Rather than defining `Body` explicitly it suffices to take as a condition that there is an isomorphism between literal bodies and values of type `Body` mediated in one direction by a judgement $b \downarrow_{\text{Body}} e_{\text{body}}$ that will come up below.

The return type, `ParseResultSE`, abbreviates a labeled sum type that distinguishes parse errors from successful parses: `ParseError` $\hookrightarrow \langle\rangle$, `SuccessE` $\hookrightarrow$ `PrExpr`.

The type abbreviated `PrExpr` classifies encodings of *proto-expressions*, $\grave{e}$ (pronounced "grave e".) The syntax of proto-expressions, defined in Figure 10, will be described when we describe proto-expansion validation in Sec. 6.6. The mapping from proto-expressions to values of type `PrExpr` is defined by the *proto-expression encoding judgement*, $\grave{e} \downarrow_{\text{PrExpr}} e$. An inverse mapping is defined by the *proto-expression decoding judgement*, $e \uparrow_{\text{PrExpr}} \grave{e}$. Again, we need only take as a condition that there is an isomorphism between values of type `PrExpr` and closed proto-expressions mediated by these judgements (see supplement.)

The third premise of Rule EE-DEF-SETSM evaluates the parse function to a value. This is not necessary, but it is the choice one would expect to make in an eager language.

The final premise of Rule EE-DEF-SETSM extends the expression TLM context, $\hat{\Psi}$, with the newly determined seTLM definition, and proceeds to assign a type, τ', and expansion, e, to $\hat{e}$. The conclusion of the rule then assigns this type and expansion to the seTLM definition as a whole.

Expression TLM contexts, $\hat{\Psi}$, are of the form $\langle \mathcal{A}; \Psi \rangle$, where $\mathcal{A}$ is a *TLM identifier expansion context* and Ψ is an *expression TLM definition context*. We distinguish TLM identifiers, $\hat{a}$, from TLM names, a, for much the same reason that we distinguish type and expression identifiers from type and expression variables: in order to allow a TLM definition to shadow a previously defined TLM definition without relying on an implicit identification convention.

An expression TLM definition context, Ψ, is a finite function mapping each TLM name $a \in \text{dom}(\Psi)$ to an *expanded seTLM definition*, $a \hookrightarrow \texttt{setlm}(\tau;\ e_{\text{parse}})$, where τ is the seTLM's type annotation, and e_{parse} is its parse function. We define $\hat{\Psi}, \hat{a} \rightsquigarrow a \hookrightarrow \texttt{setlm}(\tau;\ e_{\text{parse}})$, when $\hat{\Psi} = \langle \mathcal{A}; \Psi \rangle$, as an abbreviation of $\langle \mathcal{A} \uplus \hat{a} \rightsquigarrow a; \Psi, a \hookrightarrow \texttt{setlm}(\tau;\ e_{\text{parse}}) \rangle$.

The simple pattern TLM (spTLM) definition form operates analagously (see supplement), with the spTLM context, $\hat{\Phi}$, rather than the $\hat{\Psi}$ updated. This allows expression and pattern TLMs to use the same identifiers.

$$
\begin{array}{llll}
\text{PrTyp} & \grave{\tau} & ::= & t \mid \texttt{prparr}(\grave{\tau};\grave{\tau}) \mid \cdots \mid \texttt{splicedt}[m;n] \\
\text{PrExp} & \grave{e} & ::= & x \mid \texttt{prlam}\{\grave{\tau}\}(x.\grave{e}) \mid \texttt{prap}(\grave{e};\grave{e}) \mid \cdots \\
 & & \mid & \texttt{prmatch}[n](\grave{e};\{\grave{r}_i\}_{1\leq i\leq n}) \mid \texttt{splicede}[m;n;\grave{\tau}] \\
\text{PrRule} & \grave{r} & ::= & \texttt{prrule}(p.\grave{e}) \\
\text{PrPat} & \grave{p} & ::= & \texttt{prwildp} \mid \cdots \mid \texttt{splicedp}[m;n;\grave{\tau}]
\end{array}
$$

Fig. 10. Syntax of proto-expansions. Proto-expansion terms are ABTs identified up to alpha-equivalence.

6.5 TLM Application

The unexpanded expression form for applying an seTLM named $\hat{a}$ to a literal form with literal body b is $\hat{a}\ [\![b]\!]$. Rule EE-AP-SETSM governing this form is shown in Fig. 9.

The first premise encodes the literal body, e_{body}, which, as described above, is a value of type `Body`.

The second premise applies the parse function e_{parse} to the encoding of the literal body. If parsing succeeds, i.e. a value of the form $\texttt{inj[SuccessE]}(e_{\text{proto}})$ results from evaluation, then e_{proto} will be a value of type `PrExpr` (assuming a well-formed expression TLM context, by application of Type Safety.) We call e_{proto} the *encoding of the proto-expansion*. If the parse function produces a value labeled `ParseError`, then typed expansion fails and formally, no rule is necessary.

The third premise decodes the encoding of the proto-expansion using the judgement described in Sec. 6.4.

The fourth premise of Rule EE-AP-SETSM determines the segmentation of the proto-expansion, $\text{seg}(\grave{e})$, and ensures that it is valid with respect to b via the predicate ψ segments b, which checks that each segment in the finite set of segments ψ has non-negative length and is within bounds of b, and that the segments in ψ do not overlap.

The final premise *validates* the proto-expansion and simultaneously generates the *final expansion*, e, which appears in the conclusion of the rule. The proto-expression validation judgement is defined in the next subsection.

The typed pattern expansion rule governing pattern TLM application is analagous (see supplement).

6.6 Proto-Expansion Validation

Finally, we arrive at the crucial *proto-expansion validation judgements*, which validate the proto-expansions generated by TLMs and simultaneously generate their final expansions:

$$
\begin{array}{ll}
\Delta \vdash^{\mathbb{T}} \grave{\tau} \rightsquigarrow \tau \text{ type} & \grave{\tau} \text{ has well-formed expansion } \tau \\
\Delta\ \Gamma \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e : \tau & \grave{e} \text{ has expansion } e \text{ of type } \tau \\
\grave{p} \rightsquigarrow p : \tau \dashv\Vdash^{\mathbb{P}} \hat{\Gamma} & \grave{p} \text{ has expansion } p \text{ matching } \tau
\end{array}
$$

The purpose of the *splicing scenes* $\mathbb{T}$, $\mathbb{E}$ and $\mathbb{P}$ is to "remember" the contexts and literal body from the TLM application site (cf. the final premise of Rule EE-AP-SETSM in Fig. 9) for when validation encounters spliced terms. For example, *expression splicing scenes*, $\mathbb{E}$, are of the form $\hat{\Delta}; \hat{\Gamma}; \hat{\Psi}; \hat{\Phi}; b$.

Common Forms Most of the proto-expansion forms, including all of those elided in Fig. 10 mirror corresponding expanded forms. The rules governing proto-expansion validation for these common forms in the supplement correspondingly mirror the typing rules. Splicing scenes—$\mathbb{E}$, $\mathbb{T}$ and $\mathbb{P}$—pass opaquely through these rules, i.e. none of these rules can access the application site contexts. This maintains context independence (defined formally below.)

Notice that proto-rules, $\grave{r}$, involve expanded patterns, p, not proto-patterns, $\grave{p}$. The reason is that proto-rules appear in proto-expressions, which are generated by expression TLMs. Proto-patterns, in contrast, arise only from pattern TLMs. There is not a variable proto-pattern form, for the reasons described in Sec. 3.

References to Spliced Terms The only interesting forms are the references to spliced unexpanded types, expressions and patterns. Let us consider the rule for references to spliced unexpanded expressions:

PEV-SPLICED

$$\frac{\mathsf{parseUExp}(\mathsf{subseq}(b;m;n)) = \hat{e} \qquad \emptyset \vdash^{\hat{\Delta};\, b} \grave{\tau} \rightsquigarrow \tau\ \mathsf{type} \qquad \langle \mathcal{D}; \Delta_{\mathsf{app}} \rangle\ \langle \mathcal{G}; \Gamma_{\mathsf{app}} \rangle \vdash_{\hat{\Psi};\, \hat{\Phi}} \hat{e} \rightsquigarrow e : \tau \\ \Delta \cap \Delta_{\mathsf{app}} = \emptyset \qquad \mathrm{dom}(\Gamma) \cap \mathrm{dom}(\Gamma_{\mathsf{app}}) = \emptyset}{\Delta\ \Gamma \vdash^{\langle \mathcal{D};\Delta_{\mathsf{app}} \rangle;\, \langle \mathcal{G};\Gamma_{\mathsf{app}} \rangle;\, \hat{\Psi};\, \hat{\Phi};\, b}\ \mathtt{splicede}[m;n;\grave{\tau}] \rightsquigarrow e : \tau}$$

This first premise of this rule parses out the requested segment of the literal body, b, to produce an unexpanded expression, $\hat{e}$. The second premise performs proto-type expansion on the given type annotation, $\grave{\tau}$, producing a type, τ. The third premise then invokes type expansion on $\hat{e}$ *under the application site contexts*, $\langle \mathcal{D}; \Delta_{\mathsf{app}} \rangle$ and $\langle \mathcal{G}; \Gamma_{\mathsf{app}} \rangle$, but *not* the expansion-local contexts, Δ and Γ. The final premise requires that the application site contexts are disjoint from the expansion-local type formation context. Because proto-expansions are ABTs identified up to alpha-equivalence, we can always discharge the final premise by alpha-varying the proto-expansion. This serves to enforce capture avoidance.

The rule for references to spliced unexpanded types and patterns are fundamentally analagous (see supplement).

6.7 Metatheory

6.7.1 Typed Expansion. The first property that we are interested in is simple: that typed expansion produces a well-typed expansion. As it turns out, in order to prove this theorem, we must prove the following stronger theorem, because the proto-expression validation judgement is defined mutually inductively with the typed expansion judgement (due to splicing).

THEOREM 6.1 (TYPED EXPRESSION EXPANSION (STRONG)).

(1) If $\langle \mathcal{D}; \Delta \rangle\ \langle \mathcal{G}; \Gamma \rangle \vdash_{\langle \mathcal{A}; \Psi \rangle} \hat{e} \rightsquigarrow e : \tau$ *then* $\Delta\ \Gamma \vdash e : \tau$.

(2) If $\Delta\ \Gamma \vdash^{\langle \mathcal{D};\Delta_{app} \rangle;\, \langle \mathcal{G};\Gamma_{app} \rangle;\, \langle \mathcal{A};\Psi \rangle;\, b} \grave{e} \rightsquigarrow e : \tau$ *and* $\Delta \cap \Delta_{app} = \emptyset$ *and* $dom(\Gamma) \cap dom(\Gamma_{app}) = \emptyset$ *then* $\Delta \cup \Delta_{app}\ \Gamma \cup \Gamma_{app} \vdash e : \tau$.

The additional second clause simply states that the final expansion produced by proto-expression validation is well-typed under the combined application site and expansion-internal context (because spliced terms are distinguished only in the proto-expansion, not in the final expansion.) The combined context can only be formed if these are disjoint.

The proof proceeds by mutual rule induction and appeal to simple lemmas about type expansion and proto-type validation (see supplement). The proof is straightforward but for one issue: it is not immediately clear that the mutual induction is well-founded, because the case in the proof of part 2 for Rule PEV-SPLICED invokes part 1 of the induction hypothesis on a term that is not a sub-term of the conclusion, but rather parsed out of the literal body, b. To establish that the mutual induction is well-founded, then, we need to explicitly establish a decreasing metric. The intuition is that parsing a term of out a literal body cannot produce a bigger term than the term that contained that very literal body. The details are given in the supplemental material.

6.7.2 seTLM Reasoning Principles. The following theorem summarizes the abstract reasoning principles that programmers can rely on when applying a simple expression TLM. Informal descriptions of the labeled clauses are given inline, in gray boxes.

THEOREM 6.2 (SETLM REASONING PRINCIPLES).

If $\langle \mathcal{D}; \Delta \rangle\ \langle \mathcal{G}; \Gamma \rangle \vdash_{\hat{\Psi}} \hat{a}\ [\![b]\!] \rightsquigarrow e : \tau$ *then:*

(1) ***(Typing 1)*** $\hat{\Psi} = \hat{\Psi}', \hat{a} \rightsquigarrow a \hookrightarrow \mathtt{setlm}(\tau;\, e_{parse})$ *and* $\Delta\ \Gamma \vdash e : \tau$

The type of the expansion is consistent with the type annotation on the applied seTLM definition.

(2) $b \downarrow_{\mathsf{Body}} e_{body}$

(3) ***(Responsibility)*** $e_{parse}(e_{body}) \Downarrow \mathtt{inj[SuccessE]}(e_{proto})$
The parse function of the invoked TLM is responsible for the expansion.

(4) $e_{proto} \uparrow_{\mathsf{PrExpr}} \grave{e}$

(5) ***(Segmentation)*** $\mathsf{seg}(\grave{e})$ segments b
The segmentation determined by the proto-expansion actually segments the literal body (i.e. each segment is in-bounds and the segments are non-overlapping.)

(6) $\mathsf{seg}(\grave{e}) = \{\mathtt{splicedt}[m_i'; n_i']\}_{0 \le i < n_{ty}} \cup$
$\{\mathtt{splicede}[m_i; n_i; \grave{\tau}_i]\}_{0 \le i < n_{exp}}$

(7) ***(Typing 2)*** $\{\langle \mathcal{D}; \Delta \rangle \vdash \mathsf{parseUTyp}(\mathsf{subseq}(b; m_i'; n_i')) \rightsquigarrow \tau_i' \ \mathsf{type}\}_{0 \le i < n_{ty}}$ *and* $\{\Delta \vdash \tau_i' \ \mathsf{type}\}_{0 \le i < n_{ty}}$
Each spliced type has a well-formed expansion.

(8) ***(Typing 3)*** $\{\emptyset \vdash^{\langle \mathcal{D}; \Delta \rangle; b} \grave{\tau}_i \rightsquigarrow \tau_i \ \mathsf{type}\}_{0 \le i < n_{exp}}$ *and* $\{\Delta \vdash \tau_i \ \mathsf{type}\}_{0 \le i < n_{exp}}$
Each type annotation on a reference to a spliced expression has a well-formed expansion.

(9) ***(Typing 4)***
$\{\langle \mathcal{D}; \Delta \rangle \ \langle \mathcal{G}; \Gamma \rangle \vdash_{\hat{\Psi}} \mathsf{parseUExp}(\mathsf{subseq}(b; m_i; n_i)) \rightsquigarrow e_i : \tau_i\}_{0 \le i < n_{exp}}$ *and* $\{\Delta \ \Gamma \vdash e_i : \tau_i\}_{0 \le i < n_{exp}}$
Each spliced expression has a well-typed expansion consistent with the type annotation in the segmentation.

(10) ***(Capture Avoidance)***
$e = [\{\tau_i'/t_i\}_{0 \le i < n_{ty}}, \{e_i/x_i\}_{0 \le i < n_{exp}}]e'$ *for some variables* $\{t_i\}_{0 \le i < n_{ty}}$ *and* $\{x_i\}_{0 \le i < n_{exp}}$, *and* e'
The final expansion can be decomposed into a term with variables in place of each spliced type or expression. The expansions of these spliced types and expressions can be substituted into this term in the standard capture avoiding manner.

(11) ***(Context Independence)***
$\mathsf{fv}(e') \subset \{t_i\}_{0 \le i < n_{ty}} \cup \{x_i\}_{0 \le i < n_{exp}}$
The decomposed term makes no mention of bindings in the application site context, i.e. the only free variables are those standing for spliced terms.

Notice that we were able to state the hygiene properties (**Capture Avoidance** and **Context Independence**) without needing a notion of alpha-equivalence of source terms, as in typical formal accounts of hygiene [3, 11, 16, 29, 30, 35]. Instead, we used standard notions of capture avoiding substitution and free variables combined with the context disjointness conditions in the rules above. This is possible only because we keep track of spliced terms explicitly in the proto-expansion, rather than going straight to the final expansion.

The reasoning principles theorem for pattern TLMs is in the supplement. The key clause establishes that the hypotheses generated by the TLM application form are exactly the union of the hypothesis generated by the spliced patterns.

In the full calculus, which supports parametric TLMs, the context independence clause allows reference to the variables standing for parameters.

7 Parametric TLMs, Formally

We will now outline $\lambda^{\mathbf{P}}_{\mathbf{TLM}}$, a calculus that extends $\lambda^{\mathbf{S}}_{\mathbf{TLM}}$ with parametric TLMs. This calculus is organized, like $\lambda^{\mathbf{S}}_{\mathbf{TLM}}$, as an unexpanded language (UL) defined by typed expansion to an expanded language (XL). There is not enough space to describe $\lambda^{\mathbf{P}}_{\mathbf{TLM}}$ with the same level of detail as in Sec. 6, so we highlight only the most important concepts below. The details are in the supplement.

The XL consists of 1) module expressions, M, classified by signatures, σ; 2) constructions, c, classified by kinds, κ; and 3) expressions classified by types, which are constructions of kind Type (we use metavariables τ instead of c for types by convention.) Metavariables X ranges over module variables and u or t over construction variables. The module and construction languages are based closely on those defined by Harper [26], which in turn are based on early work by MacQueen [41, 42], subsequent work on the phase splitting interpretation of modules [27] and on using dependent singleton kinds to track type identity [12, 64], and finally on formal developments by Dreyer [15] and Lee et al. [36]. A complete account of these developments is unfortunately beyond the scope of this paper. The expression language extends the language of $\lambda^{\mathbf{S}}_{\mathbf{TLM}}$ only to allow projection out of modules.

The main conceptual difference between $\lambda^{\mathbf{S}}_{\mathbf{TLM}}$ and $\lambda^{\mathbf{P}}_{\mathbf{TLM}}$ is that $\lambda^{\mathbf{P}}_{\mathbf{TLM}}$ introduces the notion of unexpanded and expanded TLM expressions and types, as shown in Fig. 11.

$$\begin{array}{llll} \text{UMType} & \hat{\rho} ::= \hat{\tau} \mid \forall \hat{X}{:}\hat{\sigma}.\hat{\rho} & \text{MType} & \rho ::= \texttt{type}(\tau) \mid \texttt{allmods}\{\sigma\}(X.\rho) \\ \text{UMExp} & \hat{\epsilon} ::= \hat{a} \mid \Lambda \hat{X}{:}\hat{\sigma}.\hat{\epsilon} \mid \hat{\epsilon}(\hat{X}) & \text{MExp} & \epsilon ::= \texttt{defref}[a] \mid \texttt{absmod}\{\sigma\}(X.\epsilon) \mid \texttt{apmod}\{X\}(\epsilon) \end{array}$$

Fig. 11. Syntax of unexpanded and expanded TLM types and expressions in $\lambda^{\mathbf{P}}_{\mathbf{TLM}}$

The TLM type $\texttt{allmods}\{\sigma\}(X.\rho)$ classifies TLM expressions that have one module parameter matching σ. For simplicity, we formalize only module parameters. Type parameters can be expressed as module parameters having exactly one abstract type member.

The rule governing expression TLM application, reproduced below, touches all of the main ideas in $\lambda^{\mathbf{P}}_{\mathbf{TLM}}$, so we will refer to it throughout the remainder of this section.

EE-AP-PETSM

$$\frac{\begin{array}{c} \hat{\Omega} = \langle \mathcal{M}; \mathcal{D}; \mathcal{G}; \Omega_{\text{app}} \rangle \qquad \hat{\Psi} = \langle \mathcal{A}; \Psi \rangle \\ \hat{\Omega} \vdash^{\text{Exp}}_{\hat{\Psi}} \hat{\epsilon} \rightsquigarrow \epsilon \mathbin{@} \texttt{type}(\tau_{\text{final}}) \qquad \Omega_{\text{app}} \vdash^{\text{Exp}}_{\Psi} \epsilon \Downarrow \epsilon_{\text{normal}} \\ \texttt{tlmdef}(\epsilon_{\text{normal}}) = a \qquad \Psi = \Psi', a \hookrightarrow \texttt{petlm}(\rho; e_{\text{parse}}) \\ b \downarrow_{\text{Body}} e_{\text{body}} \qquad e_{\text{parse}}(e_{\text{body}}) \Downarrow \texttt{inj[SuccessE]}(e_{\text{pproto}}) \qquad e_{\text{pproto}} \uparrow_{\text{PPrExpr}} \dot{e} \\ \Omega_{\text{app}} \vdash^{\text{Exp}}_{\Psi} \dot{e} \hookrightarrow_{\epsilon_{\text{normal}}} \dot{e}\ ?\ \texttt{type}(\tau_{\text{proto}}) \dashv \omega : \Omega_{\text{params}} \\ \text{seg}(\dot{e})\ \text{segments}\ b \qquad \Omega_{\text{params}} \vdash^{\omega:\Omega_{\text{params}}; \hat{\Omega}; \hat{\Psi}; \hat{\Phi}; b} \dot{e} \rightsquigarrow e : \tau_{\text{proto}} \end{array}}{\hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{\epsilon}\ [\![b]\!] \rightsquigarrow [\omega]e : [\omega]\tau_{\text{proto}}}$$

The first two premises simply deconstruct the (unified) unexpanded context $\hat{\Omega}$ (which tracks the expansion of expression, constructor and module identifiers, as $\hat{\Delta}$ and $\hat{\Gamma}$ did in $\lambda^{\mathbf{S}}_{\mathbf{TLM}}$) and peTLM context, $\hat{\Psi}$. Next, we expand $\hat{\epsilon}$ according to straightforward unexpanded peTLM expression expansion rules. The resulting TLM expression, ϵ, must be defined at a type (i.e. no quantification over modules must remain once the literal body is encountered.)

The fourth premise performs *peTLM expression normalization*, $\Omega_{\text{app}} \vdash^{\text{Exp}}_{\Psi} \epsilon \Downarrow \epsilon_{\text{normal}}$. This is defined in terms of a structural operational semantics [57] with two stepping rules:

EPS-DYN-APMOD-SUBST-E

$$\Omega \vdash^{\text{Exp}}_{\Psi} \texttt{apmod}\{X\}(\texttt{absmod}\{\sigma\}(X'.\epsilon)) \mapsto [X/X']\epsilon$$

EPS-DYN-APMOD-STEPS-E

$$\frac{\Omega \vdash^{\text{Exp}}_{\Psi} \epsilon \mapsto \epsilon'}{\Omega \vdash^{\text{Exp}}_{\Psi} \texttt{apmod}\{X\}(\epsilon) \mapsto \texttt{apmod}\{X\}(\epsilon')}$$

Normalization eliminates parameters introduced in higher-order abbreviations, leaving only those parameter applications specified by the original TLM definition. Normal forms and progress and preservation theorems are established in the supplement.

The third row of premises looks up the applied TLM's definition by invoking a simple metafunction to extract its name, a, then looking up a within the peTLM definition context, Ψ.

The fourth row of premises 1) encodes the body as a value of the type Body; 2) applies the parse function; and 3) decodes the result, producing a *parameterized proto-expression*, $\dot{e}$. Parameterized proto-expressions, $\dot{e}$, are ABTs that serve simply to introduce the parameter bindings into an underlying proto-expression, $\grave{e}$. The syntax of parameterized proto-expressions is given below.

$$\texttt{PPrExp} \quad \dot{e} \quad ::= \quad \texttt{prexp}(\grave{e}) \mid \texttt{prbindmod}(X.\dot{e})$$

There must be one binder in $\dot{e}$ for each TLM parameter specified by a. (In Reason, we can insert these binders automatically as a convenience.)

The judgement on the fifth row of Rule EE-AP-PETSM then *deparameterizes* $\dot{e}$ by peeling away these binders to produce 1) the underlying proto-expression, $\grave{e}$, with the variables that stand for the parameters free; 2) a corresponding deparameterized type, τ_{proto}, that uses the same free variables to stand for the parameters; 3) a *substitution*, ω, that pairs the applied parameters from ϵ_{normal} with the corresponding variables generated when peeling away the binders in $\dot{e}$; and 4) a corresponding *parameter context*, Ω_{params}, that tracks the signatures of these variables. The two rules governing the proto-expression deparameterization judgement are below:

$$\frac{}{\Omega_{\text{app}} \vdash^{\text{Exp}}_{\Psi,\, a \hookrightarrow \texttt{petlm}(\rho; e_{\text{parse}})} \texttt{prexp}(\grave{e}) \twoheadrightarrow_{\texttt{defref}[a]} \grave{e} \mathbin{?} \rho \dashv \emptyset : \emptyset}$$

$$\frac{\Omega_{\text{app}} \vdash^{\text{Exp}}_{\Psi} \dot{e} \twoheadrightarrow_{\epsilon} \grave{e} \mathbin{?} \texttt{allmods}\{\sigma\}(X.\rho) \dashv \omega : \Omega \qquad X \notin \text{dom}(\Omega_{\text{app}})}{\Omega_{\text{app}} \vdash^{\text{Exp}}_{\Psi} \texttt{prbindmod}(X.\dot{e}) \twoheadrightarrow_{\texttt{apmod}\{X'\}(\epsilon)} \grave{e} \mathbin{?} \rho \dashv (\omega, X'/X) : (\Omega, X : \sigma)}$$

This judgement can be pronounced "when applying peTLM ϵ, $\dot{e}$ has deparameterization $\grave{e}$ leaving ρ with parameter substitution ω". Notice based on the second rule that every module binding in $\dot{e}$ must pair with a corresponding module parameter application. Moreover, the variables standing for parameters must not appear in Ω_{app}, i.e. $\text{dom}(\Omega_{\text{params}})$ must be disjoint from $\text{dom}(\Omega_{\text{app}})$ (this requirement can always be discharged by alpha-variation.)

The final row of premises checks that the segmentation of $\grave{e}$ is valid and performs proto-expansion validation under the parameter context, Ω_{param} (rather than the empty context, as was the case in $\lambda^{\mathbf{S}}_{\mathbf{TLM}}$.) The conclusion of the rule applies the parameter substitution, ω, to the resulting expression and the deparameterized type.

Proto-expansion validation operates conceptually as in $\lambda^{\mathbf{S}}_{\mathbf{TLM}}$. The only subtlety has to do with the type annotations on references to spliced terms. As described at the end of Sec. **??**, these annotations might refer to the parameters, so the parameter substitution, ω, which is tracked by the splicing scene, must be applied to the type annotation before proceeding recursively to expand the referenced unexpanded term. However, the spliced term itself must treat parameters parametrically, so the substitution is not applied in the conclusion of the following rule:

$$\frac{\begin{array}{c}\texttt{parseUExp}(\texttt{subseq}(b; m; n)) = \hat{e} \qquad \Omega_{\text{params}} \vdash^{\omega:\Omega_{\text{params}};\, \hat{\Omega};\, b} \grave{\tau} \rightsquigarrow \tau :: \texttt{Type} \qquad \hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \rightsquigarrow e : [\omega]\tau \\ \hat{\Omega} = \langle \mathcal{M}; \mathcal{D}; \mathcal{G}; \Omega_{\text{app}} \rangle \qquad \text{dom}(\Omega) \cap \text{dom}(\Omega_{\text{app}}) = \emptyset\end{array}}{\Omega \vdash^{\omega:\Omega_{\text{params}};\, \hat{\Omega};\, \hat{\Psi};\, \hat{\Phi};\, b} \texttt{splicede}[m; n; \grave{\tau}] \rightsquigarrow e : \tau}$$

(This is only sensible because we maintain the invariant that Ω is always an extension of Ω_{params}.)

The calculus enjoys metatheoretic properties analagous to those described in Sec. 6.7, modified to account for the presence of modules, kinds and parameterization. The following theorem establishes the abstract reasoning principles available when applying a parametric expression TLM. The clauses are directly analagous to those of Theorem 6.2, so for reasons of space we do not repeat the inline descriptions. The **Kinding** clauses can be understood by analogy to the **Typing** clauses. The details of parametric pattern TLMs (ppTLMs) are analagous (see supplement.)

Theorem 7.1 (peTLM Reasoning Principles). *If* $\hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{\epsilon}\ [\![b]\!] \rightsquigarrow e : \tau$ *then:*

(1) $\hat{\Omega} = \langle \mathcal{M}; \mathcal{D}; \mathcal{G}; \Omega_{app} \rangle$
(2) $\hat{\Psi} = \langle \mathcal{A}; \Psi \rangle$
(3) ***(Typing 1)*** $\hat{\Omega} \vdash^{\mathsf{Exp}}_{\hat{\Psi}} \hat{\epsilon} \rightsquigarrow \epsilon \ @\ \mathtt{type}(\tau)$ *and* $\Omega_{app} \vdash e : \tau$
(4) $\Omega_{app} \vdash^{\mathsf{Exp}}_{\Psi} \epsilon \Downarrow \epsilon_{normal}$
(5) $\mathsf{tlmdef}(\epsilon_{normal}) = a$
(6) $\Psi = \Psi', a \hookrightarrow \mathtt{petlm}(\rho; e_{parse})$
(7) $b \downarrow_{\mathsf{Body}} e_{body}$
(8) $e_{parse}(e_{body}) \Downarrow \mathtt{inj[SuccessE]}(e_{pproto})$
(9) $e_{pproto} \uparrow_{\mathsf{PPrExpr}} \dot{e}$
(10) $\Omega_{app} \vdash^{\mathsf{Exp}}_{\Psi} \dot{e} \hookrightarrow_{\epsilon_{normal}} \grave{e}\ ?\ \mathtt{type}(\tau_{proto}) \dashv \omega : \Omega_{params}$
(11) ***(Segmentation)*** $\mathsf{seg}(\grave{e})$ segments b
(12) $\Omega_{params} \vdash^{\omega:\Omega_{params};\hat{\Omega};\hat{\Psi};\hat{\Phi};b} \grave{e} \rightsquigarrow e' : \tau_{proto}$
(13) $e = [\omega]e'$
(14) $\tau = [\omega]\tau_{proto}$
(15) $\mathsf{seg}(\grave{e}) = \{\mathtt{splicedk}[m_i; n_i]\}_{0 \le i < n_{kind}} \cup \{\mathtt{splicedc}[m'_i; n'_i; \grave{\kappa}'_i]\}_{0 \le i < n_{con}} \cup \{\mathtt{splicede}[m''_i; n''_i; \grave{\tau}_i]\}_{0 \le i < n_{exp}}$
(16) ***(Kinding 1)*** $\{\hat{\Omega} \vdash \mathsf{parseUKind}(\mathsf{subseq}(b; m_i; n_i)) \rightsquigarrow \kappa_i\ \mathsf{kind}\}_{0 \le i < n_{kind}}$ *and* $\{\Omega_{app} \vdash \kappa_i\ \mathsf{kind}\}_{0 \le i < n_{kind}}$
(17) ***(Kinding 2)*** $\{\Omega_{params} \vdash^{\omega:\Omega_{params};\hat{\Omega};b} \grave{\kappa}'_i \rightsquigarrow \kappa'_i\ \mathsf{kind}\}_{0 \le i < n_{con}}$ *and* $\{\Omega_{app} \vdash [\omega]\kappa'_i\ \mathsf{kind}\}_{0 \le i < n_{con}}$
(18) ***(Kinding 3)*** $\{\hat{\Omega} \vdash \mathsf{parseUCon}(\mathsf{subseq}(b; m'_i; n'_i)) \rightsquigarrow c_i :: [\omega]\kappa'_i\}_{0 \le i < n_{con}}$ *and* $\{\Omega_{app} \vdash c_i :: [\omega]\kappa'_i\}_{0 \le i < n_{con}}$
(19) ***(Kinding 4)*** $\{\Omega_{params} \vdash^{\omega:\Omega_{params};\hat{\Omega};b} \grave{\tau}_i \rightsquigarrow \tau_i :: \mathsf{Type}\}_{0 \le i < n_{exp}}$ *and* $\{\Omega_{app} \vdash [\omega]\tau_i :: \mathsf{Type}\}_{0 \le i < n_{exp}}$
(20) ***(Typing 2)*** $\{\hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \mathsf{parseUExp}(\mathsf{subseq}(b; m''_i; n''_i)) \rightsquigarrow e_i : [\omega]\tau_i\}_{0 \le i < n_{exp}}$ *and* $\{\Omega_{app} \vdash e_i : [\omega]\tau_i\}_{0 \le i < n_{exp}}$
(21) ***(Capture Avoidance)*** $e = [\{\kappa_i/k_i\}_{0 \le i < n_{kind}}, \{c_i/u_i\}_{0 \le i < n_{con}}, \{e_i/x_i\}_{0 \le i < n_{exp}}, \omega]e''$ *for some* e'' *and fresh* $\{k_i\}_{0 \le i < n_{kind}}$ *and fresh* $\{u_i\}_{0 \le i < n_{con}}$ *and fresh* $\{x_i\}_{0 \le i < n_{exp}}$
(22) ***(Context Independence)*** $\mathsf{fv}(e'') \subset \{k_i\}_{0 \le i < n_{kind}} \cup \{u_i\}_{0 \le i < n_{con}} \cup \{x_i\}_{0 \le i < n_{exp}} \cup dom(\Omega_{params})$

8 Implementation

Note that even when a TLM needs only a single helper function, it must be placed into a named module. This is to avoid needing to reason about the equivalence of arbitrary expressions when determining whether two signatures are compatible.

9 Related Work

Unhygienic term rewriting systems like OCaml's PPX system and Template Haskell can be used to rewrite string literals, but these systems also do no adequately address any of these reasoning criteria. Hygienic macro systems, like those available in various Lisp-family languages and in Scala, are also unsuitable because the hygiene constraints prevent the macros from surfacing spliced expressions from literal bodies. The system of *type-specific languages* (TSLs) developed in prior work by Omar et al. [53] has made perhaps the most headway toward these ideals, but it too does not adequately maintain these reasoning criteria. Moreover, it is not clear how to integrate this approach, which was developed for a simple monomorphic language with nominal types and local type inference, into ML, which supports structural types, parameterized types, abstract types, modules, pattern matching and non-local type inference. We will say more about these existing

approaches in Sec. 9. For now, it suffices to say that after evaluating existing approaches, we found that none of them were suitable for integration into Reason.

An approach that has made some headway toward this ideal is Omar et al.'s *type-specific languages* (TSLs) [53]. A language that supports TSLs delegates control over the parsing and expansion of certain *generalized literal forms* to a parser associated with the type that the form is being checked against. So if we add TSLs to our language, our example then requires a type annotation on `y`:

```
let y : KQuery.t = `((!R)@&{&/x!/:2_!x}'!R)`
```

The parser associated with the type `KQuery.t` is statically invoked to lex, parse and expand the literal body, i.e. the sequence of characters between the outer delimiters, here `` `( `` and `` )` ``. The literal body is constrained by the context-free grammar of the language only in that nested delimiters must be matched (like comments in OCaml). TSLs are closely related to the mechanism that we will propose, so let us analyze TSLs from the perspective of the six reasoning criteria just outlined.

- ◐ **Responsibility**: The type-directed dispatch mechanism allows the client to easily determine which parser is responsible for each literal form: the parser that was associated with the type when it was defined. Clients do not need to worry about different TSLs conflicting syntactically because the context-free grammar of the language remains fixed and composition is mediated by splicing. The main limitation here is that it is impossible to define literal notation after a type has been defined, and it is also impossible to define multiple notations at a single type.
- ◐ **Segmentation**: The parser can splice expressions out of the literal body, but there is no clear way to associate each spliced term with some particular segment (i.e. subsequence) of the literal body, nor is there a guarantee that spliced terms are non-overlapping. As such, there is no way to indicate to the programmer where base language expressions are located within a literal form.
- ◐ **Capture**: The TSL mechanism as described enforces complete capture avoidance (though the formal specification given in the paper did not adequately enforce this constraint).
- ◐ **Context Dependence**: The TSL mechanism as described requires that the expansion be completely closed, so it is trivially context independent. However, this comes at a significant expressive cost: the generated expansions cannot make use of any libraries whatsoever.
- ◐ **Typing**: The expansion must necessarily be of the associated type, so abstract reasoning about the type of the expansion is straightforward. However, there is no easy way to determine the type expected for each spliced expression. In addition, the prior work considered only a monomorphic, nominally-typed language with local type inference, leaving open a number of problems that came up as we considered integrating TSLs into Reason/OCaml:
  - (a) **Structural Types**: There is no way to define literal notation at structural types, e.g. tuple and arrow types, because there is no "definition site" for such types.
  - (b) **Parameterized & Abstract Types**: There is no way to define literal notation over a type-parameterized family of types, e.g. at all types `list('a)`. Similarly, there is no way to define literal notation over a module-parameterized family of abstract types, e.g. at every abstract type defined by a module implementing the `QUEUE` signature. Parameterized and abstract type families are ubiquitous in ML-family languages.
  - (c) **Pattern Literals**: Pattern matching is ubiquitous in ML-family languages but the prior work on TSLs considered only expression literals.
  - (d) **ML-Style Type Inference**: It is not clear that a type-directed dispatch scheme can be cleanly reconciled with ML-style type inference.

9.1 Syntax Definition Systems

One approach available to library providers seeking to introduce new literal forms is to use a syntax definition system to construct a library-specific *syntax dialect*: a new syntax definition that extends the syntax definition given in the language definition with new forms, including literal forms.

There are hundreds of syntax definition systems of various design. Notable examples include grammar-oriented systems like Camlp4 [37], Copper [60, 69] and Sugar*/SoundExt [17, 18, 39, 40], as well as parser combinator systems [31]. The parsers generated by these systems can be invoked to preprocess program text in various ways, e.g. by invoking them from within a build script, by using a preprocessor-aware build system (e.g. `ocamlbuild`), or via language-integrated preprocessing directives, e.g. Racket's `#lang` directive or its reader macros [20], or the import mechanism of SugarJ [17].

The problem was described by example in Sec. 1: these systems make it difficult to reason abstractly. Let us reitierate more generally, before discussing a few systems that are exceptional along some strict subset of these dimensions:

(1) **Responsibility**: Clients using a combined syntax dialect cannot easily determine which constituent extension is responsible for a given form, whereas TLMs have explicit names which follow the usual scoping conventions.
Moreover, there can be syntactic conflicts because multiple extensions can claim responsibility for the same form. TLMs sidestep these complexities entirely because the context-free syntax of the language is fixed, and composition is via splicing rather than direct combination.
(2) **Segmentation**: Clients of a syntax dialect cannot accurately determine which segments of the program text appear directly in the desugaring. In contrast, TLM clients can inspect the inferred segmentation (communicated via secondary notation.)
(3) **Capture**: Unlike TLM clients, clients of a syntax dialect cannot be sure that spliced terms are capture avoiding.
(4) **Context Dependence**: Similarly, clients cannot be sure that the desugaring is context independent. Indeed, without a method to pass in parameters (Sec. 4), achieving strict context independence would be impractical.
(5) **Typing**: Clients of a syntax dialect cannot reason abstractly about what type a desugaring has. In contrast, TLM clients can determine the type of any expansion by referring to the parameter and type declarations on the TLM definition, and nothing else. The inferred segmentation also gives types for each spliced expression or pattern.

An extensible syntax definition system that has confronted the problem of **Responsibility** (but not the other problems) is Copper [60]. Copper integrates a modular grammar analysis that guarantees that determinism is conserved when extensions of a certain restricted class are combined. The caveat is that the constituent extensions must prefix all newly introduced forms with marking tokens drawn from disjoint sets. To be confident that the marking tokens used are disjoint, providers must base them on the domain name system or some other coordinating entity. Because the mechanism operates at the level of the context-free grammar, it is difficult for the client to define scoped abbreviations for these verbose marking tokens. TLMs can be abbreviated (Sec. 4).

Some programming languages, notably including theorem provers like Coq [44] and Agda [51], support "mixfix" notation directives [24, 49, 71]. Many of these systems enforce capture avoidance and application-site context independence [13, 24, 44, 65]. The problem is that mixfix notation requires a fixed number of sub-trees, e.g. **`if _ then _ else _`**. Coq has some limited extensions for list-like literals [44]. These systems cannot express the example literal forms from this paper, because they can have any number of spliced terms.

The work of Lorenzen and Erdweg [39, 40] introduces SoundExt, a grammar-based syntax extension system where extension providers can equip their new forms with derived typing rules. The system then attempts to automatically verify that the expansion logic (expressed using a rewrite system, rather than an arbitrary function) is sound relative to these derived rules. TLMs differ in several ways. First, as already discussed, we leave the context-free syntax fixed, so different TLMs cannot conflict. Second, SoundExt does not enforce hygiene, i.e. expansions might depend on the context and intentionally induce capture. Similarly, there is no abstract segmentation discipline. A client can only indirectly reason about binding (but not segmentation) by inspecting the derived typing rules. Unlike TLMs, SoundExt supports type-dependent expansions [40]. The trade-off is that TLMs can generate expansions, and therefore segmentations, even when the program is ill-typed. Another important distinction is that TLMs rely on proto-expansion validation, rather than verification as in SoundExt. The trade-off is that TLMs do not require that the expansion logic be written using a restricted rewriting system, nor does the system require a fully mechanized language definition. Finally, there is no clear notion of "partial application" in SoundExt or other syntax definition systems.

9.2 Term Rewriting Systems

Another approach – and the approach that TLMs are rooted in – is to leave the context-free syntax of the language fixed, and instead contextually rewrite existing literal forms.

OCaml's textual syntax now includes *preprocessor extension (ppx) points* used to identify terms that some external term rewriting preprocessor must rewrite [37]. We could mark a string literal as follows:

```
[%xml "SSTR<h1>Hello, {[first_name]}!</h1>ESTR"]
```

More than one applied preprocessor might recognize this annotation (there are, in practice, many XML/HTML libraries), so the problems of **Responsibility** comes up. It is also impossible to reason abstractly about **Segmentation**, **Capture**, **Context Dependence** and **Typing** because the code that the preprocessor generates is unconstrained.

Term-rewriting macro systems are language-integrated local term rewriting systems that require that the client explicitly apply the intended rewriting, implemented by a macro, to the term that is to be rewritten. This addresses the issue of **Responsibility**. However, unhygienic, untyped macro systems, like the earliest variants of the Lisp macro system [28], Template Haskell [62] and GHC's quasiquotation system [43] (which is based on Template Haskell), do not allow clients to reason abstractly about the remaining issues, again because the expansion that they produce is unconstrained. (It is not enough that with Template Haskell / GHC quasiquotation, the generated expansion is typechecked – to satisfy the **Typing** criterion, it must be possible to reason abstractly about *what the type of the generated expansion is*.)

Hygienic macro systems prevent, or abstractly account for [29, 30], **Capture**, and they enforce application-site **Context Independence** [3, 11, 16, 35]. The critical problem is that a standard hygiene discipline makes it impossible to repurpose string literal forms to introduce compositional literal forms at other types. Consider again our running XHTML example, which we might try to realize by applying a hygienic macro, `xml!`, to a string literal that the macro parses:

```
(xml! "SSTR<h1>Hello, {[first_name]}!</h1>ESTR")
```

The expansion of this macro will fail the check for application-site context independence because `first_name` will appear as a free variable, in no way different from any other. The hygiene mechanism for TLMs addresses this problem by explicitly distinguishing spliced segments of the literal body in the proto-expansion. This also addresses the problem of **Segmentation** – the segmentation abstractly communicates the fact that (only) `first_name` is a spliced **`string`** expression.

Much of the research on macro systems has been for languages in the LISP tradition [45] that do not have rich static type structure. The formal macro calculus studied by Herman and Wand [30] (which is not capable of expressing new literal forms, for the reasons just discussed) uses types only to encode the binding structure of the generated expansion (as discussed in Sec. 2.4.1). Research on typed *staging macro systems* like MetaML [61], MetaOCaml [34] and MacroML [21] is also not directly applicable to the problem of defining new literal forms – the syntax tree of the arguments cannot be inspected at all (staging macros are used mainly for partial evaluation and performance-related reasons.)

The Scala macro system is a hygienic macro system. Its "black box" macros support reasoning abstractly about **Typing** because type annotations constrain the macro arguments and the generated expansions, though the precise reasoning principles available are unclear because the Scala macro system has not been formally specified. The full calculus we have defined is the first detailed type-theoretic accounts of a typed, hygienic macro system of any design for an ML-like language, i.e. one with a rich static type system, support for pattern matching, type functions and ML-like modules.

Some languages, including Scala [52], build in *string splicing* (a.k.a. *string interpolation*) forms, or similar but more general *fragmentary quotation forms* [63], e.g. SML/NJ. These designate a particular delimiter to escape out into the expression language. The problem with using these together with macros as vehicles to introduce literal forms at various other types is 1) there is no "one-size-fits-all" escape delimiter, and 2) typing is problematic because every escaped term is checked against the same type. In the HTML example, we have splicing at two different types using two different delimiters. These forms also cannot appear in patterns.

This brings us back to the most closely related work, that of Omar et al. [53] on *type-specific languages* (TSLs). Like simple expression TLMs (Sec. 2), TSLs allow library providers to programmatically control the parsing of expressions of generalized literal form. With TSLs, parse functions are associated directly with nominal types and invoked according to a bidirectionally typed protocol. In contrast, TLMs are separately defined and explicitly applied. Accordingly, different TLMs can operate at the same type, and can operate at any type, including structural types. In a subsequent short paper, Omar et al. [54] suggested explicit application of simple expression TLMs also in a bidirectional typed setting [56], but this paper did not have any formal content. With TLMs, it is not necessary for the language to be bidirectionally typed (see Sec. 2.4.3 on type inference).

Perhaps most importantly, the metatheory presented by Omar et al. [53] establishes only that generated expansions are of the expected type (i.e. a variant of the Typed Expression Expansion theorem from Sec. 6.7.) It does not establish the remaining abstract reasoning principles that have been the major focus of this paper. In particular, there is no formal hygiene theorem and indeed the formal system in the paper does not correctly handle substitution or capture avoidance, issues we emphasized because they were non-obvious in Sec. 5. Moreover, the TLM does not guarantee that a valid segmentation will exist, nor associate types with segments.

Finally, the prior work did not consider pattern matching, type functions, ML-style modules, parameters or static evaluation. This paper addresses all of these.

10 Discussion

The importance of specialized notation as a "tool for thought" has long been recognized [32]. According to Whitehead, a good notation "relieves the brain of unnecessary work" and "sets it free to concentrate on more advanced problems" [8], and indeed, advances in mathematics, science and programming have often been accompanied by new notation.

Of course, this desire to "relieve the brain of unnecessary work" has motivated not only the syntax but also the semantics of languages like ML and Scala – these languages maintain a strong

type and binding discipline so that programmers, and their tools, can hold certain implementation details abstract when reasoning about program behavior. In the words of Reynolds [59], "type structure is a syntactic discipline for enforcing levels of abstraction."

Previously, these two relief mechanisms were in tension—mechanisms that allowed programmers to express new notation would obscure the type and binding structure of the program text. TLMs resolve this tension for the broad class of literal forms that generalized literal forms subsume. This class includes all of the examples enumerated in Sec. 1 (up to the choice of outermost delimiter), the case studies detailed in this paper and in the supplement, and the examples collected from the empirical study by Omar et al. [53].

Of course, not all possible literal notation will prove to be in good taste. The reasoning principles that TLMs provide, which are the primary contributions of this paper, allow clients to "reason around" poor literal designs, using principles analagous to those already familiar to programmers in languages like ML and Scala.

A correct parse function never returns an encoding of a proto-expansion that fails validation given well-typed splices, but this invariant cannot be enforced by the ML type system. Under a richer type system, the return type of the parse function itself could be refined so as to enforce this invariant intrinsically. This problem – of typed first-class typed term representations – has been studied in a variety of settings, e.g. in MetaML [61] and in the modal logic tradition [14]. Our present efforts aim to leave the semantics of OCaml unchanged.

Another direction has to do with automated refactoring. The unexpanded language does not come with context-free notions of renaming and substitution. However, given a segmentation, it should be possible to "backpatch" refactorings into literal bodies. Recent work by Pombrio et al. [58] on tracking bindings "backwards" from an expansion to the source program is likely relevant. The challenge is that the TLM's splicing logic might not be invariant to refactorings.

At several points in the paper, we allude to editor integration. However, several important questions having to do with TLM-specific syntax highlighting, incremental parsing and error recovery [22] remain to be considered, and indeed these are our biggest remaining implementation challenges.

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