Reasonably Programmable Syntax

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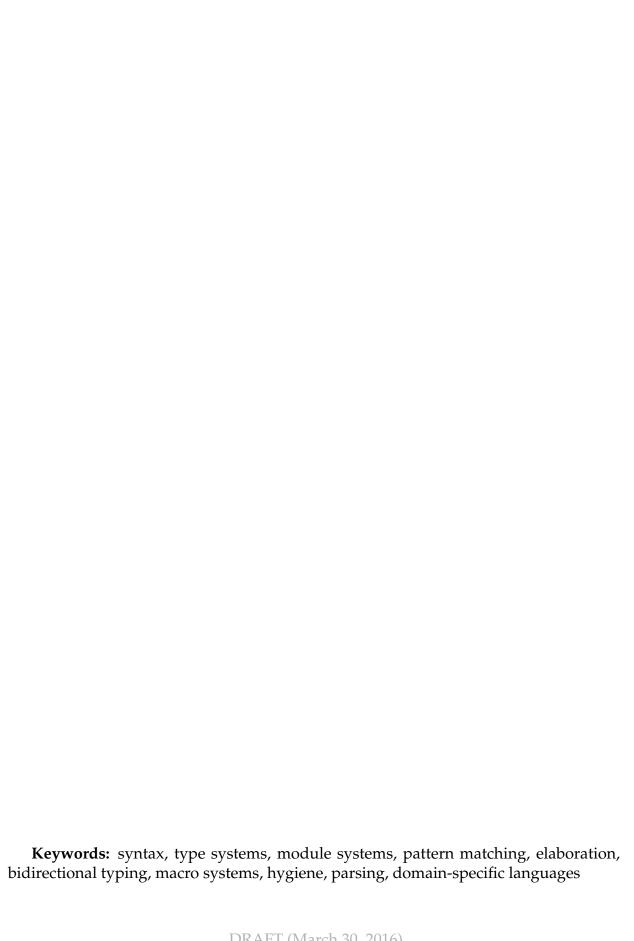Pittsburgh, PA 15213

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Abstract

When designing a human-facing textual syntax for a programming language, it is common to define "library-specific syntactic sugar", i.e. *derived forms* that decrease the cognitive cost of working with select library constructs. For example, the textual syntax of Standard ML (like that of many languages) defines derived forms for introducing and pattern matching on lists.

Defining derived forms for working with third-party library constructs, e.g. ordered finite maps, HTML elements or regular expressions, requires defining "library-specific" (a.k.a. "domain-specific") dialects of the textual syntax. Unfortunately, combining such syntax dialects, when possible at all, is not guaranteed to conserve *determinism*, i.e. syntactic ambiguities can be introduced. Moreover, it becomes difficult for programmers to reason in a disciplined manner about types and binding given only the text of a program that uses unfamiliar derived forms.

In this work, we introduce and formally define *typed syntax macros (TSMs)*, which reduce the need for *ad hoc* derived forms and library-specific syntax dialects by giving library providers the ability to programmatically control the parsing and expansion, at compile-time, of expressions and patterns of (generalized) literal form. Library clients can use any combination of TSMs in a program without needing to consider the possibility of syntactic conflicts between them, and the language validates each expansion that a TSM generates in order to maintain:

- a *type discipline*, meaning that a programmer encountering a literal form can determine its type without examining its expansion directly; and
- a *hygienic binding discipline*, meaning that
    1. the expansion cannot (accidentally or intentionally) shadow bindings that appear at the application site; and
    2. the expansion cannot make any assumptions about the application site typing context. Instead, all interactions with the application site bindings are *parametric*.

Support for partial parameter application in TSM abbreviations reduces the syntactic cost of this explicit parameter-passing style.

In short, we describe a programming language in the ML tradition with a *reasonably* programmable syntax.

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Chapter 1

Introduction

1.1 Motivation

There are many ways to draw a tree. For example, consider these three drawings:

$$1 \ / \ 2\textasciicircum3 \tag{1.1a}$$

$$\frac{1}{2^3} \tag{1.1b}$$

$$\texttt{div(num[1];pow(num[2];num[3]))} \tag{1.1c}$$

These are different drawings of a single tree – an *abstract syntax tree (AST)*, in particular, of the sort defined by the syntax chart in Figure 1.1.[1] ASTs of this sort are the expressions of a programming language, **Calc**, that can be used to perform simple arithmetic calculations with numbers. As the syntax chart suggests, every **Calc** expression can be drawn in *textual form*, like Drawing (1.1a), *stylized form*, like Drawing (1.1b), or *operational form*, like Drawing (1.1c).

The semantics of **Calc** identifies ASTs structurally, i.e. independently of how they are drawn. For instance, the semantics defines a judgement *e* val, which characterizes certain expressions as *values* (as distinct from expressions that can be further simplified.) This judgement is defined by a single inference rule, which establishes that only the number expressions are values in **Calc**:

$$\frac{}{\text{num}[n] \ \text{val}}$$

Although this inference rule is drawn using an operational form, we can nevertheless apply it to derive that 2 val, because 2 and $\text{num}[2]$ identify the same AST. In other words, "syntax doesn't matter" from the perspective of the semantics of our language.

---

[1]Some familiarity with ASTs and inductively defined judgements is preliminary to this work. See Sec. 2.1 for references and a more thorough discussion of necessary preliminaries.

Sort		Operational Form	Typeset Form	Textual Form	Description
Exp	e ::=	$\mathtt{num}[n]$	$\mathtt{n}$	$\mathtt{n}$	numbers
		$\mathtt{plus}(e;e)$	$e + e$	e + e	addition
		$\mathtt{minus}(e;e)$	$e - e$	e - e	subtraction
		$\mathtt{mult}(e;e)$	$e \times e$	e * e	multiplication
		$\mathtt{div}(e;e)$	$\frac{e}{e}$	e / e	division
		$\mathtt{pow}(e;e)$	e^e	e^e	exponentiation

Figure 1.1: Syntax of **Calc**. Metavariable n ranges over mathematical numbers (defined in some suitable manner) and $\mathtt{n}$ abbreviates the numeral forms (one for each number n, drawn in our examples in `typewriter` font.) A complete definition of the stylized and textual syntax of **Calc** would require 1) defining these numeral forms explicitly; 2) defining a form for parenthesized expressions; 3) defining the precedence and associativity of each infix operator; and 4) defining whitespace conventions. These details are not relevant to the present discussion, so we assume the usual conventions without stating them explicitly (the definition of ALGOL 60 introduced many of the conventions that persist today [47].)

That being said, most humans would nevertheless distinguish the examples drawn in stylized and textual form above, perhaps by characterizing them as more "'readable" than the example drawn in operational form. Novice programmers in particular would likely be able to grasp the intended meaning more quickly when shown the first two drawings, because they closely follow the widely-known arithmetic conventions. Similarly, the first two drawings require less effort to produce than the drawing in operational form, given typical tools. So it seems that different drawings of an AST, though semantically indistinguishable, must in fact be distinguished by the *cognitive costs* that human programmers incur when examining or producing them. We consider various operational definitions of this broad notion in Sec. 2.2.

In seeking to reduce the cognitive costs of common idioms, syntax designers often include additional *derived forms* (colloquially, "syntactic sugar") in the syntax definition. Derived forms are defined by context-independent *desugaring* (a.k.a. *rewriting* or *redrawing*) *rules*. For example, we can define a derived stylized form for calculating the square root of a **Calc** expression by including the following rule in the definition of the stylized syntax of expressions:

$$\sqrt{e} \rightarrow e^{\frac{1}{2}} \tag{1.2a}$$

Similarly, we can define a derived textual form for negating an expression by including the following rule in the definition of the textual syntax of expressions:[2]

$$-e \rightarrow \mathtt{mult}(e;\mathtt{num}[-1]) \tag{1.2b}$$

---

[2]For **Calc**, it is not necessary to restrict, for example, textual and operational forms from being interspersed – no ambiguities can arise. In richer languages, this may no longer be the case. The desugaring process must then be modified to first convert the pattern on the righthand side of a desugaring rule like Rule (1.2b) to the desired form variant before it is applied.

When we encounter a drawing of a **Calc** expression, we desugar it by first recursively desugaring its subexpressions, then, if the drawing is in derived form, applying the corresponding rule above. If this process succeeds, the resulting drawing will identify an AST, i.e. it will use the forms in Figure 1.1 (which we refer to collectively as *primitive forms*, to distinguish them from derived forms.)

Calc expressions are rather limited – they can express only simple calculations of a single type – so we should not expect to need more than a few more derived forms like these to satisfyingly capture the idioms that would arise in common human usage of **Calc**. But the study of programming languages (and logics more generally) has produced many other sorts of trees, suitable for expressing substantially richer types of computations. As humans have used these rich languages across a variety of problem domains, more idioms – and with them, more derived forms – have naturally emerged. For example, Standard ML (a "general-purpose" language in the functional tradition [27, 46]) defines derived textual forms for working with lists. In SML, [1, 2, 3, 4, 5] is the same as writing:

```
Cons(1, Cons(2, Cons(3, Cons(4, Cons(5, Nil)))))
```

if `Nil` and `Cons` are bound to the constructors of the parameterized recursive datatype `list` defined in the SML Basis library (i.e. SML's "standard library".)[3] Many other languages similarly define lists or list-like data structures in their standard libraries, and corresponding derived forms in their language definitions, anticipating that programmers are likely to use these data structures pervasively, across problem domains.

Other types of computations may be common only in more niche problem domains, which are served by third-party libraries. Language designers cannot reasonably anticipate and accomodate idioms that are specific to these libraries, so library providers (or other interested parties) sometimes define a *syntax dialect* – a new syntax definition constructed by extending an existing syntax definition with new derived forms. For example, Ur/Web extends Ur's textual syntax with derived forms for SQL queries, XHTML elements and other datatypes defined by a web programming library [10]. For example, the following program fragment shows how XHTML expressions that contain strings and other XHTML expressions can be drawn in Ur/Web:

```
val b = <xml><p>Hello, {case name of
                None => <xml>World</xml>
              | Some s => <xml>{[s]}</xml>}!</p></xml>
```

The desugaring of this form is substantially more verbose.

We will consider other examples of data structures for which derived forms can substantially lower syntactic cost in Sec. 2.3. In Sec. 2.4.4, we will describe systems like Camlp4 [40], Copper [66], SugarJ/Sugar* [14, 15] and Racket's preprocessor [20] that have lowered the costs of defining and implementing syntax dialects, and thereby contributed to their ongoing proliferation.

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[3]The desugaring actually uses unforgeable identifiers bound permanently to these constructors.

1.2 Problems with the Dialect-Oriented Approach

Some view the proliferation of syntax dialects as harmless or even as desirable, arguing that programmers can simply choose the right dialect for the job at hand [65]. However, this "dialect-oriented" approach is difficult to reconcile with the best practices of "programming in the large" [11], i.e. developing large programs "consisting of many small programs (modules), possibly written by different people" whose interactions are mediated by an expressive type and binding discipline.

1.2.1 Conservatively Combining Syntax Dialects

The first problem is that programmers cannot always "combine" different dialects when they want to use the derived forms that they define together in a single program [11].

For example, consider a syntax dialect, $\mathcal{H}$, defining derived forms for working with encodings of HTML elements, and another syntax dialect, $\mathcal{R}$, defining derived forms for working with encodings of regular expressions. In domains like bioinformatics, both HTML elements and regular expressions are common, so it would be useful to construct a "combined dialect" where all of these derived forms are defined.

For this notion of "dialect combination" to be well-defined at all, we must first have that $\mathcal{H}$ and $\mathcal{R}$ are defined using the same syntax definition system. In practice, there are many different syntax definition systems, each of which operate on subtly different classes of grammars, or that do not use grammars at all. If the dialect designers have not picked the same system, "dialect combination" remains only an informal notion.

If $\mathcal{H}$ and $\mathcal{R}$ are coincidentally defined in the same system, we must also have that this system operationalizes the notion of "dialect combination". Not all syntax definition systems do so (e.g. Racket's preprocessor is *monolithic* [20].) Clients can sometimes manually "copy-and-paste" portions of the constituent dialect definitions to construct the "combined" dialect, but this is not systematic and, in practice, quite error-prone.

If we further restrict our interest to dialects specified using a single system that does operationalize some notion of dialect combination (or equivalently one that allows programmers to systematically combine *dialect fragments*), there is likely still a problem: there is no guarantee that the combined dialect will conserve important properties that can be established about the constituent dialects in isolation (i.e. *modularly*.) In other words, establishing $P(\mathcal{H})$ and $P(\mathcal{R})$ is not sufficient to establish $P(\mathcal{H} \cup \mathcal{R})$. Clients must re-establish these properties for each combined dialect that they construct.

For example, consider two syntax dialects defined using the formalism of Camlp4: $\mathcal{D}_1$ defines derived forms for sets, and $\mathcal{D}_2$ defines derived forms for finite maps, both delimited by {| and |}.[4] Though each dialect defines a deterministic grammar, i.e. $\det(\mathcal{D}_1)$ and $\det(\mathcal{D}_2)$, when the grammars are naïvely combined by Camlp4, it is not the case that $\det(\mathcal{D}_1 \cup \mathcal{D}_2)$ (i.e. syntactic ambiguities can arise in the combined dialect.) In particular, the empty set and the empty dictionary are both drawn {||}. A third syntax dialect might use the same forms that $\mathcal{D}_2$ defines, but for ordered finite maps.

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[4]In OCaml, simple curly braces are already reserved by the language for record types and values.

Schwerdfeger and Van Wyk describe a modular analysis, implemented in Copper [66], that "nearly" guarantees that determinism is conserved when syntax dialects defined by LALR(1) grammars are combined [56], the caveat being that newly introduced derived forms must be prefixed by distinct starting tokens (which, by nature, must be verbose if they are to be predictably unique.) We will return to this requirement (and some other subtle requirements of this approach) in Sec. 2.4.4.

1.2.2 Abstract Reasoning About Derived Forms

Even putting aside the difficulties of conservatively combining syntax dialects, there are questions about how *reasonable* sprinkling library-specific derived forms throughout a large software system might be. For example, consider the perspective of a programmer attempting to comprehend (i.e. reason about) the following program fragment drawn in a syntax dialect constructed by combining SML's textual syntax with some number of other syntax dialects:

```
val a = get_a()
val w = get_w()
val x = read_data(a)
val y = {|(!R)@&{&/x!/:2_!x}'!R}|}
```

If the programmer happens to be familiar with the intentionally terse syntax of the stack-based database query processing language K, then this might pose no difficulties. If the programmer does not recognize this syntax, there is no simple, definitive protocol for answering questions like:

1. Which of the constituent dialects defined the derived form that appears on Line 4?

2. Is the character x inside this derived form parsed as a "spliced" expression, x, or parsed in some other way peculiar to this derived form?

3. If x is the spliced expression x, does it refer to the binding on the previous line? Or was that binding shadowed by an unseen binding in the desugaring?

4. If I rename w or move its binding down past the binding of y, could that possibly break the program, or change its meaning? In other words, does the desugaring rule assume that some variable identified as w is in scope?

5. What type does y have?

In contrast, when a programmer encounters, for example, a function call (e.g. the call to read_data above), in a setting where the syntax has not been extended, the analagous questions can be answered by following clear protocols that become "cognitive reflexes" after sufficient experience with the language:

1. The language's syntax definition determines how read_data(a) should be parsed.

2. According to the syntax definition, read_data and a are parsed as variables.

3. The variable a on Line 3 can only refer to the binding of a on Line 1.

4. read_data(a) does not mention w, so the binding of w can be renamed, moved or removed without changing the meaning of Line 3.

5. The type of `x` can be determined to be `B` by first determining that the type of `read_data` is `A -> B` for some `A`, and then checking that a has type `A`. Nothing else needs to be known about the values of `read_data` or `a`.

In other words, languages like ML are structured so as to provide programmers with powerful abstract reasoning principles. In the words of Reynolds [54]:

Type structure is a syntactic discipline for enforcing levels of abstraction.

Syntax dialects, on the other hand, do not maintain *syntactic abstraction* – if the desugaring of a program is held abstract, programmers can no longer reason about types and binding in the usual disciplined manner.

1.3   Contributions

Our aim in this work is to introduce primitive language constructs that reduce the need for syntax dialects. In particular, we introduce **typed syntax macros**, or **TSMs**. TSMs are applied to trees of generalized literal form, which are definitively delimited but whose bodies are initially left unparsed, to programmatically control their expansion to trees that do not contain such forms (i.e. trees in *expanded form*.)

Syntactic conflicts between TSMs are impossible by construction, because the syntax of the language is never directly modified – only contextually repurposed. Expansion occurs simultaneously with typing, in a phase that we call *typed expansion*. As such, the semantics can take the type and binding structure of the surrounding program into account to impose constraints that serve to ensure that clients can maintain a reasonable type discipline and binding discipline, i.e. that questions like those above can be answered without examining the expansion itself. More specifically, TSMs maintain a *hygienic binding discipline*, meaning that the class of perverse expansions that Questions 4 and 5 above were concerned with are disallowed entirely. We will, of course, make these notions more technically precise as we continue.

We introduce TSMs first in a simple language of expressions and types in Chapter 3, then add support for pattern matching in Chapter 4. In Chapter 5, we add an ML-style module system, and define *parametric TSMs*, i.e. TSMs that take type and module parameters. This allows the expansions that parametric TSMs generate to interact with the surrounding context in a controlled manner, and also allows library providers to define TSMs that operate not just at a single type, but uniformly over a type- and module-parameterized family of types.

In the chapters just mentioned, we make a simplifying assumption: that each TSM definition is fully self-contained, i.e. that TSM definitions do not need access to libraries. This allows us to focus on the fundamental contributions of this work, but it is, of course, an unrealistic assumption in practice. We relax this assumption in Chapter 6.

In Chapter 7, we show how library clients can contextually designate, for any type, a privileged TSM at that type, and then rely on a bidirectional type system to invoke that TSM implicitly. This method of *TSM implicits* can reduce the cognitive cost of an idiom to very nearly the same extent that a special-purpose dialect can, while still maintaining the reasoning principles just summarized.

As vehicles for this work, we define a small programming language in each of the chapters just mentioned, each building upon the previous one. All formal claims made in this work involve these small languages.

In examples, we assume a full-scale functional language called VerseML.[5] VerseML is the language of Chapter 7 extended with a few additional conveniences that are commonly found in other functional languages and semantically orthogonal to TSMs (e.g. higher-rank polymorphism [13], signature abbreviations, and a number of standard derived forms that are not library-specific, e.g. for curried functions.) We do not formally define these features to avoid unnecessarily complicating our presentation with details that are not essential to the ideas presented herein – our purpose is only to introduce new language constructs, not to define a full-scale programming language. As such, all examples involving VerseML should be understood to be informal motivating material for the subsequent formal material.

Thesis Statement

In summary, this work defends the following statement:

> A functional programming language can give library providers the ability to programmatically control the parsing and expansion of expressions and patterns of generalized literal form while maintaining a type discipline, a hygienic binding discipline and modular reasoning principles.

1.4   Disclaimers

Before we continue, it may be prudent to explicitly acknowledge that eliminating the need for dialects would indeed be asking for too much: certain syntax design decisions are fundamentally incompatible with others or require coordination across a language design. We aim only to decrease the need for syntax dialects in this work. We summarize some of the situations that we explicitly do not consider when we discuss future work in Sec. 8.3.

It may also be useful to explicitly acknowledge that library providers could use TSMs to define constructs that are in "poor taste". We expect that in practice, VerseML will come with a standard library defining an expertly curated collection of TSMs, as well as guidelines for advanced users regarding when it would be sensible to define their own TSMs (following the example of languages that support operator overloading or *ad hoc* polymorphism using type classes [12, 26], which also have some potential for "abuse" or "overuse".) The majority of programmers would not be expected to define TSMs themselves.

---

[5]We distinguish VerseML from Wyvern, which is the language described in our prior publications about some of the work that we will describe, because Wyvern is a group effort evolving independently.

Chapter 2

Background

2.1 Preliminaries

This work is rooted in the tradition of full-scale functional languages with non-trivial type structure like Standard ML, OCaml and Haskell (as might have been obvious from Chapter 1.) Familiarity with basic concepts in these languages, e.g. variables, types, polymorphic and recursive functions, tuples, records, recursive datatypes and structural pattern matching, is assumed throughout this work. Readers who are not familiar with these concepts are encouraged to consult the early chapters of an introductory text like Harper's *Programming in Standard ML* [27] (a working draft can be found online.) We will briefly discuss integration of TSMs into languages from other design traditions in Sec. 8.6.3.

Chapter 5 and some of the motivating examples given below, consider questions of integration with an ML-style module system, so readers with experience in a language without such a module system (e.g. Haskell) are also advised to consult the relevant chapters in *Programming in Standard ML* [27] as needed.

The formal systems that we will construct are defined within the metatheoretic framework of type theory. More specifically, we assume that abstract syntax trees (ASTs), abstract binding trees (ABTs, which enrich ASTs with the notions of binding and scope), substitution, renaming, alpha-equivalence, structural induction and rule induction are defined as described in Harper's *Practical Foundations for Programming Languages, Second Edition* (*PFPL*) [29] (a working draft can be found online), except as otherwise stated. Familiarity with other formal accounts of type systems, e.g. Pierce's *Types and Programming Languages* (*TAPL*) [53], should also suffice. This document is organized so as to be readable even if the sections defining formal systems are skipped entirely (although much precision will, of course, be lost).

2.2 Cognitive Cost

The notion of *cognitive cost* is central to the motivation of this work (though it does not enter into any of the technical material directly.) Ultimately, this notion must be under-

stood intuitively, relating as it does to the complexities of the human mind. Cognitive cost is also fundamentally a *subjective* notion. Nevertheless, it is often useful to operationalize this notion in various ways.

One way to operationalize cognitive cost is by reducing it to a simpler notion of *syntactic cost*, measured, for example, by counting glyphs. This is often a satisfying proxy for cognitive cost, in that smaller drawings are usually easier to comprehend and produce. For example, the drawing [1, 2, 3, 4, 5] has lower syntactic cost than its desugaring, as discussed in the previous chapter. There is a limit to this approximation, of course. For example, one might argue that the syntax for K shown in the previous chapter, though it has low syntactic cost, nevertheless has high cognitive cost, at least until one is well-experienced with its use. In other words, the relationship between syntactic cost and cognitive cost depends on the subject's progression on a *learning curve*.

A related property of interest to human programmers in practice is *edit cost*, measured relative to a program editor as the minimum number of primitive edit actions that must be performed to produce a drawing. For example, when using a text editor (as most professional programmers do), drawings in textual forms typically have lower edit cost, as measured by the minimum number of keystrokes necessary to draw them, than operational or stylized forms (in fact, some stylized forms may not be expressible at all with a text editor alone, in which case they can be understood to have infinite edit cost.) Edit cost can be modeled using, for example, *keystroke-level models* (KLMs) as introduced by Card, Moran and Newell [9].

One can also analyze cognitive cost using qualitative methods. For example, Green's *Cognitive Dimensions of Notations* [23, 24] and Pane and Myers' *Usability Issues* [52] (both of which synthesized much of the earlier work in the area) are popular heuristic frameworks. Using Green's cognitive dimensions framework, for example, it could be argued that the derived list forms described in Chapter 1 *map more closely* to other notations used for sequences of elements (e.g. in typeset mathematics, or on a physical notepad.) They also make the elements of the list more clearly *visible*, in that the identifier Cons is not interspersed throughout the term. Finally, the derived forms have lower *viscosity* because adding a new item to the middle of a list drawn in derived form requires only a local edit, whereas for a list drawn in primitive form, one needs also to add a closing parenthesis to the end of the term.

Finally, one might consider cognitive cost comparatively using empirical methods, e.g. by conducting randomized control trials to compare forms with respect to task completion time or error rate (or the derivative of error rate, which can be plotted as a learning curve.) Stefik et al. have performed many such studies, mainly focused on novice programmers (summarized, along with other studies, in [59].)

2.3 Motivating Definitions

In this section, we will define a number of VerseML data structures that we will use as examples in the remainder of this document. This section also serves to introduce the textual syntax and semantics of VerseML.

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2.3.1 Lists

In VerseML, lists are defined by the following *type function*:

```
type 'a list = rec(self => [Nil + Cons of 'a * self])
```

Here, `list` is a type function binding its type argument to the type variable `'a`. Types are applied to type functions in prefix position. For example, the type of integer lists is `int list`, which is equal to:

```
rec(self => [Nil + Cons of int * self])
```

This is a *recursive type*, binding the type variable `self` for use as a *self reference*. Values of a recursive type `T` are introduced as (`fold` e) : `T`, where the type of e must be the *unfolding* of `T`, computed by substituting `T` for `self`. For example, the unfolding of `int list` is the following type:

```
[Nil + Cons of int * int list]
```

This *labeled sum type* specifies two *variants*. One, labeled `Nil`, specifies a trivial payload, and the other, labeled `Cons`, specifies a payload of product type, `int * int list`. Values of a labeled sum type `T` are introduced as (`inj`[lbl] e) : `T`, where lbl is one of the labels that `T` specifies, and e is a value of the corresponding payload type. Values of product type are introduced as tuples.

When introducing a value of recursive type or labeled sum type, the type ascription can be omitted if it can be inferred. For example, the empty integer list is:

```
val nil_int : int list = fold(inj[Nil] ())
```

and a list containing a single integer, 0, is:

```
val one_int : int list = fold(inj[Cons] (0, nil_int))
```

To lower syntactic cost, we can define the following polymorphic values, which we call the *list constructors*, that apply the fold and injection operators for us:

```
val Nil : 'a list = fold(inj[Nil] ())
fun Cons (x : 'a * 'a list) : 'a list =>
  fold(inj[Cons] x)
```

In fact, VerseML generates constructors like these automatically. Using these constructors, we can write `nil_int` and `one_int` at lower syntactic cost as follows:

```
val nil_int : int list = Nil
val one_int : int list = Cons (0, Nil)
```

Values of recursive type, labeled sum type and product type are deconstructed by pattern matching. For example, we can write the polymorphic map function, which constructs a new list by applying a function to each element of a given list, as follows:

```
fun map (f : 'a -> 'b) (xs : 'a list) : 'b list =>
  match xs with
  | fold(inj[Nil] ()) => Nil
  | fold(inj[Cons] (y, ys)) => Cons (f y, map f ys)
  end
```

Readers who are not familiar with structural pattern matching may wish to consult the introduction to Chapter 4 for a somewhat more detailed description.

To lower syntactic cost, VerseML automatically inserts the folds, injections and trivial arguments into patterns of constructor form:

```
fun map (f : 'a -> 'b) (xs : 'a list) : 'b list =>
  match xs with
  | Nil => Nil
  | Cons (y, ys) => Cons (f y, map f ys)
  end
```

To avoid syntactic ambiguity, patterns can only bind variables whose identifiers begin with a lowercase letter, and the labels in a labeled sum type must begin with an uppercase letter.

We package the definitions above (as well as some other standard utility functions) into a module List : LIST, where LIST is a signature that is defined as follows:

```
signature LIST =
sig
  type 'a list = rec(self => [Nil + Cons of 'a * self])
  val Nil : 'a list
  val Cons : 'a * 'a list -> 'a list
  val map : ('a -> 'b) -> 'a list -> 'b list
  val append : 'a list -> 'a list -> 'a list
  (* ... *)
end
```

The textual syntax of VerseML does not build in derived list forms. In subsequent chapters, we will show how they can be approximated as a mode of use of TSMs.

2.3.2 Regular Expressions

Regular expressions, or *regexes*, are the expressions of a structural query language that operates over strings [62]. Regexes are particularly prevalent in domains like natural language processing and bioinformatics.

Recursive Sums   One simple way to encode regular expressions within VerseML is by defining a recursive labeled sum type:

```
type rx = rec(rx => [Empty + Str of string + Seq of rx * rx +
                    Or of rx * rx + Star of rx])
```

Figure 2.1: Definition of the recursive labeled sum type rx

Assuming the automatically generated constructors as in Sec. 2.3.1, we can construct a regex that matches the strings "A", "T", "G" or "C" (i.e. DNA bases) as follows:

```
Or(Str "A", Or(Str "T", Or(Str "G", Str "C")))
```

```
fun is_dna_rx(r : rx) : boolean =>
  match r with
  | Str "A" => True
  | Str "T" => True
  | Str "G" => True
  | Str "C" => True
  | Seq (r1, r2) => (is_dna_rx r1) andalso (is_dna_rx r2)
  | Or  (r1, r2) => (is_dna_rx r1) andalso (is_dna_rx r2)
  | Star(r') => is_dna_rx r'
  | _ => False
  end
```

Figure 2.2: Pattern matching in VerseML

Given a value of type rx, we can deconstruct it by pattern matching. For example, the function is_dna_rx defined in Figure 2.2 detects regular expressions that match DNA sequences:

Abstract Types   Encoding regexes as values of type rx is straightforward, but there are reasons why one might not wish to expose this encoding to clients directly.

First, regexes are usually identified up to their reduction to a normal form. For example, Seq(Empty, Str "A") has normal form Str("A"). It can be useful for regexes with the same normal form to be indistinguishable from the perspective of client code.

Second, it can be useful for performance reasons to maintain additional data alongside regexes (e.g. a corresponding finite automaton), while hiding this "implementation detail" from clients. In fact, there may be many ways to represent regular expression patterns, each with different performance trade-offs. This suggests that we would like to provide clients with a choice of implementations.

For these reasons, another approach in VerseML, as in SML, is to abstract over the choice of representation using the module system. In particular, we define the signature RX, shown in Figure 2.3, where the representation type of patterns, t, is held abstract.

Clients of any module R that has been sealed against RX, e.g. R1 and R2 in Figure 2.3, manipulate regexes as values of type R.t using the interface specified by this signature. Because the identity of R.t is held abstract by the signature, the only way for clients to construct values of this type is by using the operations that RX specifies. Consequently, R can choose (and evolve) its internal representation freely and need only establish any internal representation invariants locally. As such, R.t is an *abstract data type* (ADT) [41].

For example, we can construct a regex matching DNA bases as follows:

```
R.Or(R.Str "A", R.Or(R.Str "T", R.Or (R.Str "G", R.Str "C")))
```

Because the representation type is held abstract, we cannot interrogate the structure of a value r : R.t directly. Instead, the signature specifies a function unfold_norm that produces the *normal unfolding*[1] of the given regex, i.e. a value of type t u that exposes

---

[1]This sense of the word "unfolding" is conceptually related to, but technically distinct, from the sense

```
type 'a u = [UEmpty + UStr of string + USeq of 'a * 'a +
            UOr of 'a * 'a + UStar of 'a]

signature RX =
sig
  type t

  val Empty : t
  val Str : string -> t
  val Seq : t * t -> t
  val Or : t * t -> t
  val Star : t -> t

  val unfold_norm : t -> t u
end

module R1 : Rx = struct (* ... *) end
module R2 : Rx = struct (* ... *) end
```

Figure 2.3: Definition of the RX signature and two example implementations.

the outermost form of the regex after it has been normalized (the details of regex normalization are not important for our purposes.) The client can pattern match over the normal unfolding in the familiar manner:

```
fun is_dna_rx'(r : R.t) : boolean =>
  match R.unfold_norm r with
  | UStr "A" => True
  | UStr "T" => True
  | UStr "G" => True
  | UStr "C" => True
  | USeq (r1, r2) => (is_dna_rx' r1) andalso (is_dna_rx' r2)
  | UOr (r1, r2) => (is_dna_rx' r1) andalso (is_dna_rx' r2)
  | UStar r' => is_dna_rx' r'
  | _ => False
  end
```

The function above only needed to examine the outermost structure of r. However, in general, we may want to pattern match more deeply into a regex. To do so, we must first construct the n-deep unfolding of r, where n is the depth of the deepest pattern. For example, it is easy to define a function `unfold_norm2 : R.t -> R.t u u` in terms of `R.unfold_norm` that allows pattern matching to depth 2.[2]

Alternatively, we might completely unfold a value of type t to produce a value of the type rx that was defined previously. We can compute the complete unfolding (also called the *view*) by applying a utility function `view : R.t -> rx`, also defined in terms of `R.unfold_norm`.

in which it is used for recursive types discussed above.

[2]Defining an unfolding *generic* in n is a more subtle problem that is beyond the scope of this work.

```
functor RXUtil(R : RX) =
struct
  fun unfold_norm2(r : R.t) : R.t u u =>
    (* ... *)

  fun view(r : R.t) : rx =>
    match R.unfold_norm r with
    | UEmpty => Empty
    | UStr s => Str s
    | USeq (r1, r2) => Seq (view r1, view r2)
    | UOr (r1, r2) => Or (view r1, view r2)
    | UStar r => Star (view r)
    end

  (* ... *)
end
```

Figure 2.4: The definition of `RXUtil`.

Typically, utility functions like `unfold_norm2` and `view` are defined in a functor (i.e. a function at the level of modules) like `RXUtil` in Figure 2.4, so that they need only be defined once, rather than separately for each module `R : RX`. The client can instantiate the functor by applying it to any such module as follows:

```
module RU = RXUtil(R)
```

2.4 Existing Approaches

We will now review existing approaches that library providers seeking to decrease the cognitive cost of idioms involving definitions like those above might consider.

2.4.1 Standard Abstraction Mechanisms

The simplest approach is to capture idioms using the standard abstraction mechanisms of our language, i.e. bound values, functions and modules.

We already saw examples of this approach above. For example, we defined the "list constructors", which capture the idioms of list construction:

```
val Nil : 'a list = fold(inj[Nil] ())
fun Cons (x : 'a * 'a list) : 'a list =>
  fold(inj[Cons] x)
```

Such definitions are common enough that VerseML generates them automatically.

We also defined a utility functor for regexes, `RXUtil`, in Figure 2.4. As more idioms involving regexes arise, we can capture them by adding additional definitions to this functor (or some other such functor, if we cannot modify `RXUtil` itself.) For example, we can define a value that matches digits:

```
val digit = R.Or(R.Str "0", R.Or(R.Str "1", ...))
```

Similarly, we can define a function `repeat : R.t -> int -> R.t` that computes a regex by sequentially repeating the given regex a given number of times. Using these auxiliary definitions, a client can define a regex that matches social security numbers as follows:

```
val dash = R.Str "-"
val repeat_d = RU.repeat RU.digit
val ssn = R.Seq(repeat_d 3, R.Seq(dash, R.Seq(repeat_d 2,
                R.Seq(dash, repeat_d 4))))
```

One limitation of this approach is that it provides no way to capture idioms at the level of patterns. Expressions and patterns are two different sorts of trees, and a pattern must be fully determined statically, whereas variables stand for dynamic values and functions perform dynamic computation.

Another limitation is that this approach does not give library providers control over form. For example, we cannot "approximate" SML-style derived list forms using only auxiliary definitions (otherwise, SML would not have defined derived list forms.) Similarly, consider the textual syntax for regexes defined in the POSIX standard [4] (which data suggests that most professional programmers are familiar with [49].) Using this syntax, the regex that matches DNA bases is drawn:

```
A|T|G|C
```

Similarly, the regex that matches SSNs just discussed is drawn:

```
\d\d\d-\d\d-\d\d\d\d
```

or

```
\d{3}-\d{2}-\d{4}
```

These drawings have lower syntactic cost than those above, and programmers familiar with the POSIX syntax for regular expressions would likely agree that these drawings have lower cognitive cost as well.

String Parsing

We might attempt to approximate the POSIX syntax for regexes by defining a function `parse : string -> R.t` in `RXUtil` that parses a string representation of a POSIX regex form to produce a regular expression value, or raises an exception if the input is malformed with respect to the POSIX specification.

Using `RU.parse`, we can construct the regex matching DNA bases as follows:

```
RU.parse "A|T|G|C"
```

This approach is imperfect for several reasons:

1. First, there are syntactic conflicts between standard string escape sequences and standard regex escape sequences. For example, the following is not a well-formed drawing in Standard ML:

```
val ssn = RU.parse "\d\d\d-\d\d-\d\d\d\d" (* ERROR *)
```

In practice, a parser would provide the client with an error message like:[3]

```
error: illegal escape character
```

because \d is not a valid string escape character (even though it is a valid regex escape character.) In a small lab study, we observed that this class of error confused even experienced programmers if they had not used regexes recently [49].

One workaround – escaping all backslashes – nearly doubles syntactic cost:

```
val ssn = RU.parse "\\d\\d\\d-\\d\\d-\\d\\d\\d\\d"
```

Some languages, anticipating this use of string literals, build in alternative string forms that leave escape sequences uninterpreted. For example, OCaml supports alternative string literals like {rx|\d\d\d-\d\d-\d\d\d\d|rx}.

2. The next problem is that dynamic string parsing only decreases the syntactic cost of complete regexes. Regexes constructed compositionally do not benefit from this technique. For example, the function below constructs a regex from a string, `name`, and another regex, `ssn`:

```
fun lookup_rx(name : string) =>
  R.Seq(R.Str name, R.Seq(R.Str ": ", ssn))
```

Figure 2.5: Compositional construction of regexes.

We cannot use `RU.parse` to redraw this equivalently, but at lower syntactic cost.

(We will describe a syntax dialect that does capture such idioms in Sec. 2.4.4. In particular, we will compare Figure 2.5 with Figure 2.7.)

This is also why dynamic string parsing is irrelevant for capturing idioms like list construction – list expressions contain sub-expressions.

3. Using strings to introduce regexes also creates a *cognitive hazard* for programmers who are coincidentally working with other data of type `string`. For example, consider the following seemingly "more readable definition of `lookup_rx`", where the infix operator ^ means string concatenation:

```
fun lookup_rx_insecure(name : string) =>
  RU.parse (name ^ {rx|: \d\d\d-\d\d-\d\d\d\d|rx})
```

or equivalently, given the regex `ssn` as above and an auxiliary function `RU.to_string` that can generate its string representation:

```
fun lookup_rx_insecure(name : string) =>
  RU.parse (name ^ ": " ^ (RU.to_string ssn))
```

Both `lookup_rx` and `lookup_rx_insecure` have the same type, `string -> R.t`, and behave identically at many inputs, particularly the "typical" inputs (i.e. alphabetic

---

[3]This is the error message that `javac` produces. When compiling an analagous expression using SML of New Jersey (SML/NJ), we encounter a more confusing error message: `Error:  unclosed string`.

names.) It is only when `lookup_rx_insecure` is applied to a string that corresponds to a regex that matches more than just that string that it behaves incorrectly.

In applications that query sensitive data, mistakes like this lead to *injection attacks*, which are among the most common and catastrophic security threats today [5].

This problem is, fundamentally, attributable to the programmer making a mistake in a misguided effort to decrease syntactic cost. However, the availability of a better approach for reducing syntactic cost would serve to make this class of mistakes less prevalent [7]. Given that our design philosophy is explicitly concerned with issues of cognitive cost, it is natural to also consider common cognitive hazards.

4. The final problem is that regex parsing does not occur until the call to `RU.parse` is dynamically evaluated. For example, the malformed string encoding of a regex in the program fragment below will only trigger an exception when this expression is evaluated during the full moon:

```
match moon_phase with
Full => RU.parse "(GC" | _ => (* ... *)
end
```

Malformed string encodings of regexes can sometimes be discovered by testing, though empirical data gathered from large open source projects suggests that there remain many malformed regexes that are not detected by a project's test suite "in the wild" [58].

One workaround is for the programmer to lift all such calls where the argument is a string literal out to the top level of the program, so that the exception is raised every time the program is evaluated. This subtly changes the performance profile of the program, and there is a cognitive penalty associated with moving the description of a regex away from its use site, but for complete regexes, this might be an acceptable trade-off.

Problems like these arise whenever a library provider attempts to deploy dynamic string parsing as a solution to the problem of high syntactic cost. The reason is that syntactic cost is a property of a drawing of a program, so trying to address it by drawing a different program requires establishing that the alternative program is equivalent to the program that the client would write if syntactic cost was not a consideration (which is, at worst, an ill-posed problem, and at best, a rather difficult problem.)

This is not to say that one should simply refrain from defining or applying a function like `RU.parse`. There are valid uses of string parsing that are not motivated by the desire to decrease syntactic cost, e.g. when parsing regular expressions received as dynamic input to the program.

2.4.2 Quotation Parsing

Some syntax dialects of ML, e.g. a syntax dialect available via a compiler flag in SML/NJ [2], define *quotation literals*: derived forms for expressions of type `'a frag list`, where `'a frag` is defined as follows (using SML's datatype declaration syntax):

```
datatype 'a frag = QUOTE of 'a | ANTIQUOTE of string
```
Quotation literals are delimited by backticks, e.g. `A|T|G|C` is the same as writing
`[QUOTE "A|T|G|C"]`. Expressions of variable or parenthesized form that appear prefixed
by a caret in the body of a quotation literal are parsed out and appear wrapped in the
`ANTIQUOTE` constructor, e.g. `GC^(dna_rx)GC` is the same as writing

```
[QUOTE "GC", ANTIQUOTE dna_rx, QUOTE "GC"]
```
It is possible to define a function `qparse : R.t frag list -> R.t` in `RXUtil` that allows
clients to dynamically construct regexes from fragment lists like these.

Similarly, it is possible to define a function `qparse : 'a frag list -> 'a list` in the
`List` module that allows clients to use quotation literals to construct lists:

```
List.qparse '[^(x + y), ^y, ^z]'
```

As with dynamic string parsing, parsing occurs dynamically. We cannot use the trick
of lifting all calls to `qparse` to the top level of our program, because the arguments are
no longer string literals. At best, we can lift these calls into the earliest possible "stage"
of evaluation. Parse errors are only detected once this stage is entered, and the dynamic
cost of parsing is incurred each time this stage is entered. For example, `List.qparse` is
called n times below, where n is the length of `input`:

```
List.map (x => List.qparse '[^x, ^(2 * x)]') input
```

Another problem is that the antiquote character, i.e. the caret, is fixed *a priori*. This
is problematic for regexes, for example, because the caret has a different meaning in the
POSIX standard (and in practice, appears quite often.)

All antiquoted values must be of the same type. To support both spliced regexes and
spliced strings, for example, we need to define an auxiliary datatype in `RXUtil` and the
client needs to apply it in each antiquotation. These "marking constructors" increase
syntactic cost. For example, we we would need to write `lookup_rx` as follows:

```
fun lookup_rx(string : name) =>
  RU.qparse' '^(RU.QS name): ^(RU.QR reading)'
```

Similarly, we would need to use marking constructors to support quoted lists where
the tail is explicitly given by the client (following OCaml's revised syntax [40]):

```
List.qparse '[^(List.V x), ^(List.V y) :: ^(List.VS zs)]'
```

Finally, quotation parsing, like the other approaches considered thusfar, helps only
with the problem of abbreviating expressions. It provides no solution to the problem of
abbreviating patterns.

VerseML does not build in quotation literals.[4]

2.4.3   Infix Function Application

To gain more precise control over form, a library provider, or another interested party,
might also consider defining a syntax dialect using a syntax definition system.

---

[4]In fact, quotation syntax can be expressed using parametric TSMs, which are the topic of Chapter 5,
though we will leave the details as an exercise for the reader. As such, even in the odd situation where
parsing fragment lists drawn using quotations is the right solution, VerseML is a suitable language.

```
1  infix 5 ::
2  infix 6 <*>
3  infix 4 <|>
4
5  functor RXOps(R : RX) =
6  struct
7    module RU = RXUtil(R)
8    val op:: = R.Seq
9    val op<*> = RU.repeat
10   val op<|> = R.Or
11 end
```

Figure 2.6: Fixity declarations and related bindings for RX.

The simplest syntax definition systems allow programmers to designate identifiers as infix operators. For example, SML builds in such a system. In SML, we can define :: as a right-associative infix operator at precedence level 5 as follows:

```
infixr 5 ::
```

In a context where the variable **op::** is bound to the list constructor that we identified as Cons earlier, we can construct a list with x as its head and y :: zs as its tail as follows:

```
x :: y :: zs
```

In the SML Basis library, the list datatype is defined such that the constructor that we labeled Cons is instead labeled **op::**. As such, clients can use infix :: in patterns as well. Binding the variable **op::** to Cons is not equivalent in this regard (again, because variables do not stand for patterns.)

Figure 2.6 shows how our regex library might define several such fixity declarations, together with a functor RXOps that binds the corresponding operator variables to the appropriate functions. Assuming that the library packaging system has brought the fixity declarations and the definition of RXOps from Figure 2.6 into scope, the client can instantiate RXOps with their choice of module R : RX and then **open** this instantiated module (also called a *structure* in SML) to make these bindings visible.

```
structure ROps = RXOps(R)
open ROps
```

From there, the client can draw the examples discussed earlier equivalently as follows:

```
val dna = (R.Str "A") <|> (R.Str "T") <|> (R.Str "G") <|>
          (R.Str "C")
val ssn = (RU.digit)<*>3 :: (RU.digit)<*>2 :: (RU.digit)<*>4
fun lookup_rx(name : string) =>
  (Str name) :: (Str ": ") :: ssn
```

This demonstrates the two main limitations of this approach.

First, it grants only limited control over form – we cannot express the POSIX forms in this way, only *ad hoc* (and in this case, rather poor) approximations thereof. Moreover, the infix operators declared in Figure 2.6 cannot be used for pattern matching, again because variables only stand for values.

The second limitation has to do with the possibility of syntactic conflicts. Notice that both the list library and the regex library have declared :: an infix operator, but with different associativity. They also export different bindings for **op::**. As such, clients cannot use both in the same scope. There is no mechanism that allows a client to explicitly qualify the use of an infix operator as referring to the fixity declaration from one or the other library – fixity declarations are purely syntactic (i.e. they only influence parsing) and are not exported from modules or otherwise integrated into the binding structure of ML (note that libraries are extralinguistic packaging constructs, distinct from modules.)

Formally, each fixity declaration induces a dialect of the subset of SML's textual syntax that does not allow the declared identifier to be used in prefix position. When two such dialects are combined, the resulting dialect is not necessarily a dialect of both of the constituent dialects (the fixity declaration that is used depends on the order in which they are combined.)

Due to these limitations, VerseML does not inherit this mechanism from SML (the infix operators that are available in VerseML, like ^ for string concatenation, have a fixed precedence and associativity.)

2.4.4   More General Syntax Definition Systems

Fixity declarations give their users no direct control over desugaring – the desugaring is always the corresponding function application. More general syntax definition systems give users more direct control over the desugaring process. In this section, we will first review these more general syntax definition systems, then give two examples of syntax dialects, followed by a discussion of the difficulties that programmers can encounter if they try to adopt a dialect-oriented approach "in the large" (as a follow-up to the problems already introduced in Sec. 1.2.)

Mixfix Notational Definitions

Some languages, including many popular type-theoretic proof assistants like Coq [45], support "mixfix" notational definitions. These generalize SML-style fixity declarations, in that newly defined forms can contain any fixed number of sub-terms (rather than just two). Their desugarings are determined by a user-defined rewriting rule. Griffin gives a formal account of such systems [25], and Taha and Johann incorporate Griffin's account into a more general and well-behaved account of *staged notational definitions* [60].

Although more expressive than fixity declarations, these syntax definition systems also do not allow us to express POSIX regex syntax as-is. The issue is fundamentally that these systems only give programmers the ability to extend the syntax of the existing sorts of trees, e.g. expressions and patterns. They do not give programmers the ability to define new sorts of trees, with their own distinct syntax. For example, we cannot define a sort for regular expressions, where sequences of characters are not recognized as identifiers but rather as regex character sequences.

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Grammar-Based Syntax Definition Systems

Many syntax definition systems are oriented around *formal grammars* [33]. Formal grammars have been studied since at least the time of Pãnini, who developed a grammar for Sanskrit in or around the 4th century BCE [36].

Context-free grammars (CFGs) were first used to define the textual syntax of a major programming language – Algol 60 – by Backus [47]. Since then, countless other syntax definition systems oriented around CFGs have emerged – we will only summarize the systems that are particularly relevant to this work. In these systems a syntax definition consists of a CFG (perhaps from some restricted class of CFGs) equipped with various auxiliary definitions (e.g. a lexer/scanner definition in some systems) and logic for computing an output value (e.g. a tree) based on the determined form of the input.

Perhaps the most established systems in the ML ecosystem are ML-Lex + ML-Yacc, which is distributed with SML/NJ [61], and Camlp4, which was (until recently) integrated into the OCaml system (in recent releases of the OCaml system, it has been deprecated in favor of a simpler system, ppx, that we discuss in the next section) [40]. In these systems, output is computed by ML functions that appear associated with each production in the grammar (these functions are referred to as the *semantic actions*.)

The *syntax definition formalism (SDF)* [31] is a syntactic formalism for describing CFGs. SDF is used by a number of syntax definition systems, e.g. the Spoofax "language workbench" [37]. These systems commonly use Stratego, a rule-based rewriting language, as the language that output logic is written in [64]. SugarJ is an extension of Java that allows programmers to define and combine fragments of SDF+Stratego-based syntax definitions directly from within the program text [15]. SugarHaskell is a similar system based on Haskell [17] and Sugar* simplifies the task of defining such extensions of other languages [14]. SoundExt and SugarFOmega add the requirement that new derived forms must come equipped with corresponding derived typing rules [42]. The system must be able to verify that the rewrite logic is sound with respect to these rules (their verification system is written in, and defers to the facilities of, PLT-Redex [19].) SoundX generalizes this idea to support other base languages, and also adds the ability to define type-dependent rewritings [43].

Copper implements a CFG-based syntax definition system that uses a context-aware scanner, which is useful for reasons that we will discuss shortly [66]. Silver is an *attribute grammar* system based on Copper [63]. Attribute grammars are used to compositionally define output logic that requires information that is not local to each form [38].

Some other syntax definition systems are instead oriented around *parsing expression grammars* (PEGs) [21]. PEGs are similar to CFGs, distinguished mainly in that they are deterministic by construction (by allowing only for explicitly prioritized choice between alternative parses.)

Parser Combinator Systems

Parser combinator systems define a functional interface for defining a parser, together with various functions that generate new parsers from given parsers and other values (these

```
val ssn = /\d\d\d-\d\d-\d\d\d\d/
fun lookup_rx(name : string) => /@name: %ssn/
```

Figure 2.7: Derived regex expression forms in $\mathcal{V}_{\mathbf{rx}}$

functions are referred to as the *parser combinators*) [34]. In some cases, it is acceptable to take an expression that involves parser combinators as definitional (as opposed to the usual view, where a parser is an *implementation* of a separate syntax definition.)

For example, Hutton describes a system where parsers are functions of some type in the following type family:

```
type 'char 'tree parser = 'char list -> ('tree*'char list) list
```

i.e., a parser is a function that takes a list of (abstract) characters and returns a list of valid parses, each of which consists of an (abstract) tree and a list of the characters that were not consumed. An input is malformed if this function returns the empty list, and it is ambiguous if this function returns more than one parse. A deterministic parser is one that never returns more than one parse. The parser combinator `alt`, declared

```
val alt : 'c 't parser -> 'c 't parser -> 'c 't parser
```

combines the two input parsers by applying them both to the input and appending the lists that they return, i.e. it defines non-deterministic choice.

Various alternative formulations of this concept that better control dynamic cost or have other useful properties have also been described. For example, Hutton and Meijer describe a parser combinator in monadic style [35]. Okasaki has described an alternative design that uses continuations to control cost [48].

Example 1: $\mathcal{V}_{\mathbf{rx}}$

Let us now consider a dialect of VerseML's textual syntax called $\mathcal{V}_{\mathbf{rx}}$, defined using some suitably expressive syntax definition system like those just described, that builds in derived forms related to the recursive type `rx` that was defined in Figure 2.1.[5]

$\mathcal{V}_{\mathbf{rx}}$ extends the syntax of expressions with *derived regex literals*, which are delimited by forward slashes, e.g.:

```
/A|T|G|C/
```

The desugaring of this form is equivalent to the following, assuming an environment where the variables `Or` and `Str` stand for the corresponding regex constructors:

```
Or(Str "A", Or (Str "T", Or (Str "G", Str "C")))
```

It is not reasonable to assume that `Or` and `Str` are bound appropriately at every use site. In order to maintain *context independence*, the desugaring applies the explicit **fold** and **inj** operators as discussed in Sec. 2.3.1.[6]

---

[5]Technically, $\mathcal{V}_{\mathbf{rx}}$ must be a dialect of the textual syntax of VerseML's expansion language; cf. Chap. 3.

[6]In SML, where datatypes are generative and the constructors values are necessary, we would need to use a technique like the one described for $\mathcal{V}_{\mathrm{RX}}$, next.

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```
fun is_dna_rx(r : rx) : boolean =>
  match r with
  | /A/ => True
  | /T/ => True
  | /G/ => True
  | /C/ => True
  | /%(r1)%(r2)/ => (is_dna_rx r1) andalso (is_dna_rx r2)
  | /%(r1)|%(r2)/ => (is_dna_rx r1) andalso (is_dna_rx r2)
  | /%(r)*/ => is_dna_rx r'
  | _ => False
  end
```

Figure 2.8: Derived regex pattern forms in $\mathcal{V}_{\mathrm{rx}}$

```
fun is_dna_rx'(r : R.t) : boolean =>
  match R.unfold_norm r with
  | /A/ => True
  | /T/ => True
  | /G/ => True
  | /C/ => True
  | /%(r1)%(r2)/ => (is_dna_rx' r1) andalso (is_dna_rx' r2)
  | /%(r1)|%(r2)/ => (is_dna_rx' r1) andalso (is_dna_rx' r2)
  | /%(r)*/ => is_dna_rx r'
  | _ => False
  end
```

Figure 2.9: Derived pattern forms in $\mathcal{V}_{\mathrm{RX}}$

$\mathcal{V}_{\mathrm{rx}}$ also supports regex literals that contain subexpressions, to capture the idioms that arise when constructing regexes compositionally. For example, the definition of lookup_rx in Figure 2.7 is equivalent to the definition of lookup_rx that was given in Figure 2.5, i.e. it constructs a regex from a string, name, and another regex, ssn. The prefix @ followed by the identifier name causes the expression name to appear in the desugaring as if it was wrapped in the Str constructor, and the prefix % followed by the identifier ssn causes ssn to appear in the desugaring directly. We refer to the subexpressions that appear inside literal forms as *spliced subexpressions*.

To splice in an expression that is not of variable form, e.g. a function application, we must delimit it with parentheses:

```
/@(capitalize name): %(ssn)/
```

Finally, $\mathcal{V}_{\mathrm{rx}}$ extends the syntax of patterns with analagous *derived regex pattern literals*. For example, the definition of is_dna_rx in Figure 2.8 is equivalent to the definition of is_dna_rx that was given in Figure 2.2. Notice that the variables bound by the patterns in Figure 2.8 appear inside *spliced sub-patterns*.

Example 2: $\mathcal{V}_{\mathbf{RX}}$

In Sec. 2.3.2, we gave a more sophisticated formulation of our regex library organized around the signature RX defined in Figure 2.3. Let us define another dialect of VerseML's textual syntax called $\mathcal{V}_{\mathrm{RX}}$ that defines derived forms whose desugarings involve modules that implement RX. For this to work in a context-independent manner, these forms must take the particular module that is to appear in the desugaring as a spliced term. For example, in the following program fragment, the module R is "passed into" each derived form for use in its desugaring as a spliced module:

```
val ssn = R./\d\d\d-\d\d\d\d-\d\d\d/
fun lookup_rx'(name : string) => R./@name: %ssn/
```

The desugaring of the body of `lookup_rx'` is:

```
R.Seq(R.Str(name), R.Seq(R.Str ": ", ssn))
```

This is context-independent because the constructors are explicitly qualified (i.e. Seq and Str are *field labels* here, not variables.) The only variables are R, name and ssn, all of which were given by the client at the application site.

Recall that RX specifies a function `unfold_norm : t -> t u` for computing the normal unfolding of the given regex, which is a value of type R.t u. $\mathcal{V}_{\mathrm{RX}}$ defines derived forms for patterns matching values of types in the type family 'a u. These appear in the definition of `is_dna_rx'` given in Figure 2.9.

Combining $\mathcal{V}_{\mathbf{rx}}$ and $\mathcal{V}_{\mathbf{RX}}$

Notice that the derived regex pattern forms that appear in Figure 2.9 are identical to those that appear in Figure 2.8. Their desugarings are, however, different. In particular, the patterns in Figure 2.9 match values of type R.t u, whereas the patterns in Figure 2.8 match values of type rx.

These two examples were written in different syntax dialects. However, it would be useful to have derived forms for values of type rx available even when we are working with a value of a type R.t, because we have defined a function `view : R.t -> rx` in RXUtil. This brings us to the first of the two main problems with the dialect-oriented approach, already described in Chapter 1: there is no good way to conservatively combine $\mathcal{V}_{\mathrm{rx}}$ and $\mathcal{V}_{\mathrm{RX}}$. In particular, any such "combined dialect" will either fail to conserve determinism, or the combined dialect will not be a dialect of both of the constituent dialects, i.e. some of the forms from one dialect will "shadow" the overlapping forms from the other dialect (depending on the order in which they were combined [21].)

In response to this problem, Schwerdfeger and Van Wyk have developed a "nearly" modular analysis that accepts only deterministic extensions of a base LALR(1) grammar where all new forms must start with a "marking" terminal symbol and obey certain other constraints related to the follow sets of the base grammar's non-terminals [56]. By relying on a context-aware scanner (a feature of Copper [66]) to transfer control when the marking terminals are seen, extensions of a base grammar that pass this analysis and specify distinct marking terminals can be combined without introducing conflict.

```
fun is_dna_rx'(r : R.t) : boolean =>
  match R.unfold_norm r with
  | $cmu_edu_comar_rx $u/A/ => True
  | $cmu_edu_comar_rx $u/T/ => True
  | $cmu_edu_comar_rx $u/G/ => True
  | $cmu_edu_comar_rx $u/C/ => True
  (* and so on *)
  | _ => False
  end
```

Figure 2.10: Using URL-based marking tokens to avoid syntactic conflicts.

The analysis is "nearly" modular in that only a relatively simple "combine-time" check that the set of marking terminals is disjoint is necessary.

For the two dialects just considered, these conditions are not satisfied. If we modify the grammar of $\mathcal{V}_{RX}$ so that, for example, the regex literal forms are marked with $r and the regex unfolding forms were marked with $u, the analysis will accept both grammars, and the combine-time disjointness check will pass, solving our immediate problem at only a small cost. However, a conflict could still arise later when a client combines these extensions with another extension that also uses the marking terminals $r, $u or /.

The solution given in [56] is 1) to allow for the grammar's name to be used as an additional syntactic prefix when a conflict arises, and 2) to adopt a naming convention for grammars based on the Internet domain name system (or some similar coordinating system.) Figure 2.9 shows how a client needs to draw is_dna_rx' when a conflict arises. Clearly, this drawing has substantially higher syntactic cost than the drawing in Figure 2.9. Moreover, there is no simple way for clients to selectively control this cost by defining scoped abbreviations for marking tokens or grammar names (as one does for types, modules or values that are exported from deeply nested modules) because this mechanism is only syntactic, i.e. agnostic to the binding structure of the base language.

TODO: mention this? http://www.ccs.neu.edu/home/ejs/papers/tfp12-island.pdf

Abstract Reasoning About Derived Forms

In addition to the difficulties of conservatively combining syntax dialects, there are a number of other difficulties related to the fact that there is often no useful notion of syntactic abstraction that a programmer can rely on to reason about an unfamiliar derived form. The programmer may need to examine the desugaring, the desugaring logic or even the definitions of all of the constituent dialects, to definitively answer the questions given in Sec. 1.2.2. These questions were stated relative to a particular example involving the query processing language K. Here, we generalize from that example to develop an informal classification of the difficulties that programmers might encounter in analagous situations. In each case, we will discuss exceptional systems where these difficulties are ameliorated or avoided.

1. **Search:** It is not always straightforward to determine which constituent dialect is responsible for any particular derived form.

   The system implemented by Copper [56] is an exception, in that the marking terminal (and the grammar name, if necessary) allows clients to search across the constituent dialect definitions for the corresponding declaration without needing to understand any of them deeply.

2. **Segmentation:** It is not always possible to segment a derived form such that each segment consists either of a spliced base language term (which we have drawn in black in the examples in this document) or a sequence of characters that are parsed otherwise (which we have drawn in color.) When it is possible, determining the segmentation is not always straightforward.

   For example, consider a production in a grammar that looks like this:

   ```
   start <- "%(" verseml_exp ")"
   ```

   The name of the non-terminal `verseml_exp` suggests that it will match any VerseML expression, but it is not certain that this is the case. Moreover, even if we know that this non-terminal matches VerseML expressions, it is not certain that the output logic will insert that expression as-is into the desugaring – it may instead only examine its form, or transform it in some way (in which case highlighting it as a spliced expression might be misleading.)

   Systems that support the generation of editor plug-ins, such as Spoofax [37] and Sugarclipse for SugarJ [16], can generate syntax coloring logic from an annotated grammar definition, which often give programmers some indication of where a spliced term occurs. However, there is no definitive information about segmentation in how the editor displays the derived form. (Moreover, these editor plug-ins can themselves conflict, even if the syntax itself is deterministic.)

3. **Shadowing:** The desugaring of a derived form might place spliced terms under binders. These binders are not visible in the program text, but can shadow those that are. This obscures the binding structure of the program.

   For derived forms that desugar to module-level definitions (e.g. to one or more **val** definitions), a desugaring might introduce locally-scoped bindings and, simultaneously, exported module components that are similarly invisible in the text. This can cause both local shadowing as well as non-local shadowing if a client **open**s the module into scope.

   In most cases, shadowing is inadvertent. For example, a desugaring might bind an intermediate value to some temporary variable, `tmp`. This can cause problems at use sites where `tmp` is bound. If the types of the two bindings are incompatible, the problem will be caught statically. Otherwise, it will cause unanticipated dynamic behavior. It is easy to miss this problem in testing.

   In some syntax dialects, shadowing is by design. For example, in (Sugar)Haskell, **do** notation for monadic values introduces a new binding construct [17]. For programmers who are familiar with **do** notation, this can be useful. But when a pro-

grammer encounters an unfamiliar form, this forces them to determine whether it similarly is designed as a new binding construct. A simple grammar provides no information about shadowing.

In most systems, it is possible for dialect providers to generate identifiers that are guaranteed to be fresh at the use site. If dialect providers are disciplined about using this mechanism, they can prevent such conflicts. However, this is awkward and most systems provide not guarantee that the dialect provider maintained this freshness discipline [18].

To enforce a prohibition on shadowing, the system must be integrated into or otherwise made aware of the binding structure of the language. For example, some of the language-integrated mixfix systems discussed above, e.g. Coq's notation system [45], enforce a prohibition on shadowing by alpha-renaming desugarings as necessary. Erdweg et al. have developed a formalism for directly describing the "binding structure" of program text, as well as contextual transformations that use these descriptions to rename the identifiers that appear in a desugaring (and more generally, a rewriting) to avoid shadowing [18, 55].[7]

4. **Context Dependence:** If the desugaring of a derived form assumes that certain identifiers are bound at the use site (e.g. to particular values, or to values of some particular type), we refer to the desugaring as being *context dependent*.

Context dependent desugarings take control over naming away from clients. Moreover, it is difficult to determine the assumptions that a desugaring is making. As such, it is difficult to reason about whether renaming an identifier or moving a binding is a meaning-preserving transformation.

In our examples above, we maintained context independence as a "courtesy" by explicitly applying the **fold** and **inj** operators, or by taking the module for use in the desugaring as a "syntactic argument".

To enforce context independence, the system must be aware of binding structure and have some way to distinguish those subterms of a desugaring that originate in the text at the use site (which should have access to bindings at the use site) from those that do not (which should only have access to bindings internal to the desugaring.) For example, language-integrated mixfix systems, e.g. Coq's notation system, use a simple rewriting system to compute desugarings, so they satisfy these requirements and can enforce context independence. Coq gives desugarings access only to the bindings visible where the notation was defined.

More flexible systems where desugarings are computed functionally, or language-external systems that have no understanding of binding structure, do not enforce context independence.

5. **Typing:** Finally, and perhaps most importantly, it is not always clear what type an expression drawn in derived form has, or what type of value that a pattern drawn

---

[7]These papers refer to this property as "capture avoidance". We use the term "shadowing" rather than "capture" because "capture" has several incompatible meanings in the literature.

in derived form matches.

Similarly, it is not always straightforward to determine what type a spliced expression has, or what type of value that a spliced pattern matches.

SoundExt/SugarFomega [43] and SoundX [55] allow dialect providers to define derived typing rules alongside derived forms and desugaring rules. These systems automatically verify that the desugaring rules are sound with respect to these derived typing rules. This ensures that type errors are never reported in terms of the desugaring (which is the stated goal of their work). However, this helps only to a limited extent in answering the questions just given. In particular, the programmer must construct a derivation using the derived typing rules introduced by all of the constituent dialects, then examine this derivation to answer questions about the type of the desugaring and the spliced terms within it.

Even for relatively simple base languages, like System $\mathbf{F}_\omega$, understanding a typing derivation requires significantly more expertise than programmers usually need.[8] For languages like ML, the judgement forms are substantially more complex.

As discussed in Sec. 1.2, languages with a rich type and binding structure are designed to minimize or eliminate the cognitive cost of answering analagous questions. These reasoning principles are central to the claim that such languages are suitable for "programming in the large" [11].

Due to the problems of syntactic conflict and the reasoning difficulties enumerated above, the textual syntax of VerseML cannot be modified or extended from within.

2.4.5 Rewriting Systems

Another approach is to leave the textual syntax of the language fixed, but repurpose it for novel ends using a *term rewriting system*.

Language-External Term Rewriting

Language-external rewriting systems operate as *preprocessors*, transforming well-formed program fragments to produce other well-formed program fragments.

For example, one could define a preprocessor that rewrites every string literal that is followed by the comment `(*rx*)` to the corresponding expression (or pattern) of type `rx`, following the approach discussed in the previous section. For example, the following expression would be rewritten to a regex expression, with `dna` treated as a spliced subexpression as described in the previous section:

```
"GC%(dna)GC" (*rx*)
```

OCaml 4.02 introduced *extension points* into its textual syntax [**?**]. Extension points serve as markers for the benefit of a preprocessor. They are less *ad hoc* than comments, in

---

[8]At CMU, we teach ML to all first-year students (in 15-150.) However, understanding a judgemental specification of a language like System $\mathbf{F}_\omega$ involves skills that are taught only to some third and fourth year students (in 15-312.)

that each extension point is associated with a single term in a well-defined way, and the compiler gives an error if any extension points remain after preprocessing is complete. For example, in the following program fragment,

```
let%lwt (x, y) = f in x + y
```

the `%lwt` annotation on the let expression causes a preprocessor distributed with `Lwt`, a lightweight threading library, to rewrite this fragment to:

```
Lwt.bind f (fun (x, y) -> x + y)
```

The OCaml system is distributed with a library called `ppx_tools` that simplifies the task of writing preprocessors that operate on terms annotated with extension points.

These systems have nearly all of the same conceptual problems as dialect-oriented systems do. TODO: elaborate on this

Macro Systems

Macro systems allow programmers to designate functions that operate over term encodings as macros, and then apply these macros directly to terms. TODO: search is solved. delimitation is solved. conflict is solved.

The LISP macro system [30] is perhaps the most prominent example of such a system. Early variants of this system suffered from the problem of hygiene described earlier TODO: elaborate, but later variants, notably in the Scheme dialect of LISP, brought support for enforcing hygiene [39].

In languages with a richer static type discipline, variants of macros that restrict rewriting to a particular type and perform the rewriting statically have also been studied [22, 32] and integrated into languages, e.g. MacroML [22] and Scala [8]. TODO: actually MacroML isn't a term rewriting system; just a staging system. write about this. TODO: typing is sort of solved...

The most immediate problem with using these for our example is that we are not aware of any such statically-typed macro system that integrates cleanly with an ML-style module system. In other words, macros cannot be parameterized by modules. However, let us imagine such a macro system. We could use it to repurpose string syntax as follows:

```
let val ssn = rx R {rx|\d\d\d-\d\d-\d\d\d\d|rx}
```

The definition of the macro `rx` might look like this:

```
1  macro rx[Q : RX](e) at Q.t {
2    static fun f(e : Exp) : Exp => case(e) {
3        StrLit(s) => (* regex parser here *)
4      | BinOp(Caret, e1, e2) => `Q.Seq(Q.Str(%e1), %(f e2))`
5      | BinOp(Plus, e1, e2) => `Q.Seq(%(f e1), %(f e2))`
6      | _ => raise Error
7    }
8  }
```

Here, `rx` is a macro parameterized by a module matching `rx` (we identify it as `Q` to emphasize that there is nothing special about the identifier `R`) and taking a single argument, identified as `e`. The macro specifies a type annotation, **at** `Q.t`, which imposes the

constraint that the expansion the macro statically generates must be of type `Q.t` for the provided parameter `Q`. This expansion is generated by a *static function* that examines the syntax tree of the provided argument (syntax trees are of a type `Exp` defined in the standard library; cf. SML/NJ's visible compiler [3]). If it is a string literal, as in the example above, it statically parses the literal body to generate an expansion (the details of the parser, elided on line 3, would be entirely standard). By parsing the string statically, we avoid the problems of dynamic string parsing for statically-known patterns.

For patterns that are constructed compositionally, we need to get more creative. For example, we might repurpose the infix operators that are normally used for other purposes to support string and pattern splicing, e.g. as follows:

```
fun example_using_macro(name : string) =>
  rx R (name ^ ": " + ssn)
```

The binary operator `^` is repurposed to indicate a spliced string and `+` is repurposed to indicate a spliced pattern. The logic for handling these forms can be seen above on lines 4 and 5, respectively. We assume that there is derived syntax available at the type `Exp`, i.e. *quasiquotation* syntax as in Lisp [6] and Scala [57], here delimited by backticks and using the prefix `%` to indicate a spliced value (i.e. unquote).

Having to creatively repurpose existing forms in this way limits the effect a library provider can have on cognitive cost (particularly when it would be desirable to express conventions that are quite different from the conventions adopted by the language). It also can create confusion for readers expecting parenthesized expressions to behave in a consistent manner. However, this approach is preferable to direct syntax extension because it avoids many of the problems discussed above: there cannot be syntactic conflicts (because the syntax is not extended at all), we can define macro abbreviations because macros are integrated into the language, there is a hygiene discipline that guarantees that the expansion will not capture variables inadvertently, and by using a typed macro system, programmers need not examine the expansion to know what type the expansion produced by a macro must have.

Part I

Simple TSMs

DRAFT (March 30, 2016)

Chapter 3

Unparameterized Expression TSMs (ueTSMs)

We now introduce a new primitive – the **typed syntax macro** (TSM). TSMs, like term-rewriting macros (Sec. 2.4.5), generate expansions. Unlike term-rewriting macros, TSMs are applied to unparsed *generalized literal forms*, which gives them substantially more syntactic flexibility. This chapter considers perhaps the simplest manifestation of TSMs: **unparameterized expression TSMs** (ueTSMs), which generate expressions of a single specified type. We will consider unparameterized pattern TSMs (upTSMs) in Chapter 4 and parameterized TSMs (pTSMs) in Chapter 5.

3.1 Expression TSMs By Example

We begin in this section with a "tutorial-style" introduction to ueTSMs in VerseML. In particular, we discuss a ueTSM for constructing values of the recursive labeled sum type Rx that was defined in Figure 2.1. We then formally specify ueTSMs with a reduced calculus, miniVerse$_{\mathbf{UE}}$, in Sec. 3.2.

3.1.1 Usage

In the following concrete VerseML expression, we apply a TSM named `$rx` to the *generalized literal form* `/A|T|G|C/`:

```
$rx /A|T|G|C/
```

Generalized literal forms are left unparsed when concrete expressions are first parsed. It is only during the subsequent *typed expansion* process that the TSM parses the *body* of the provided literal form, i.e. the characters between forward slashes in blue here, to generate a *candidate expansion*. The language then *validates* the candidate expansion according to criteria that we will establish in Sec. 3.1.4. If candidate expansion validation succeeds, the language generates the *final expansion* (or more concisely, simply the *expansion*) of the expression. The program will behave as if the expression above has been replaced by its expansion. The expansion of the expression above, written concretely, is:

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```
1  'body cannot contain an apostrophe'
2  `body cannot contain a backtick`
3  [body cannot contain unmatched square brackets]
4  {body cannot contain an unmatched curly brace}
5  /body cannot contain a forward slash/
6  \body cannot contain a backslash\
```

Figure 3.1: Generalized literal forms available for use in VerseML's concrete syntax. The characters in blue indicate where the literal bodies are located within each form. In this figure, each line describes how the literal body is constrained by the form shown on that line. The Wyvern language specifies additional forms, including whitespace-delimited forms [50] and multipart forms [51], but for simplicity we leave these out of VerseML.

```
  Or(Str "A", Or(Str "T", Or(Str "G", Str "C")))
```

A number of literal forms, shown in Figure 3.1, are available in VerseML's concrete syntax. Any literal form can be used with any TSM, e.g. we could have equivalently written the example above as `$rx` `A|T|G|C` (in fact, this would be convenient if we had wanted to express a regex containing forward slashes but not backticks). TSMs have access only to the literal bodies. Because TSMs do not extend the concrete syntax of the language directly, there cannot be syntactic conflicts between TSMs.

3.1.2 Definition

Let us now take the perspective of the library provider. The definition of the TSM `$rx` shown being applied above has the following form:

```
syntax $rx at Rx {
  static fn(body : Body) : CEExp ParseResult =>
    (* regex literal parser here *)
}
```

This TSM definition first names the TSM. TSM names must begin with the dollar symbol ($) to clearly distinguish them from variables (and thereby clearly distinguish TSM application from function application). This is inspired by a similar convention enforced by the Rust macro system [1].

The TSM definition then specifies a *type annotation*, **at** Rx, and a *parse function* within curly braces. The parse function is a *static function* responsible for parsing the literal body when the TSM is applied to generate an encoding of the candidate expansion, or an indication of an error if one cannot be generated (e.g. when the body is ill-formed according to the syntactic specification that the TSM implements). Static functions are functions that are applied during the typed expansion process. For this reason, they do not have access to surrounding variable bindings (because those variables stand in for dynamic values). For now, let us simply assume that static functions are closed (we discuss introducing a distinct class of static bindings so that static values can be shared between TSM definitions in Sec. 6.2).

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```
type IndexRange = {startIndex : nat, endIndex : nat}

type 'a ParseResult = Success of 'a
                    | ParseError of {
                        msg : string, loc : IndexRange
                      }
```

Figure 3.2: Definitions of `IndexRange` and `ParseResult` in the VerseML prelude.

```
type CETyp = TyVar of var_t
           | Arrow of CETyp * CETyp
           | (* ... *)
           | Spliced of IndexRange

type CEExp = Var of var_t
           | Fn of var_t * CETyp * CEExp
           | App of CEExp * CEExp
           | (* ... *)
           | Spliced of IndexRange
```

Figure 3.3: Abbreviated definitions `CETyp` and `CEExp` in the VerseML prelude. We assume some suitable type `var_t` exists, not shown.

The parse function must have type `Body -> CEExp ParseResult`. These types are defined in the VerseML *prelude*, which is a collection of definitions available ambiently. The input type, `Body`, gives the parse function access to the body of the provided literal form. For our purposes, it suffices to define `Body` as an abbreviation for the `string` type:

```
type Body = string
```

The output type, `CEExp ParseResult`, is a labeled sum type that distinguishes between successful parses and parse errors. The parameterized type `'a ParseResult` is defined in Figure 3.2.

If parsing succeeds, the parse function returns a value of the form $\text{Success}(e_{\text{cand}})$, where e_{cand} is the *encoding of the candidate expansion*. Encodings of candidate expansions are, for expression TSMs, values of the type `CEExp` defined in Figure 3.3 (in Chapter 5, we will introduce pattern TSMs, which generate patterns rather than expressions; this is why `ParseResult` is defined as a parameterized type). Expressions can mention types, so we also need to define a type `CETyp` in Figure 3.3. We discuss the constructors labeled `Spliced` in Sec. 3.1.3; the remaining constructors (some of which are elided for concision) encode the abstract syntax of VerseML expressions and types. To decrease the syntactic cost of working with the types defined in Figure 3.3, the prelude provides *quasiquotation syntax* at these types, which is itself implemented using TSMs. We will discuss these TSMs in more detail in Sec. 6.1. The definitions in Figure 3.3 are recursive labeled sum types to simplify our exposition, but we could have chosen alternative encodings of terms, e.g. based on abstract binding trees [29], with only minor modification to the semantics.

If the parse function determines that a candidate expansion cannot be generated, i.e. there is a parse error in the literal body, it returns a value labeled by `ParseError`. It must provide an error message and indicate the location of the error within the body of the literal form as a value of type `IndexRange`, also defined in Figure 3.2. This information can be used by VerseML compilers when reporting errors to the programmer.

3.1.3 Splicing

To support splicing syntax, like that described in Sec. 2.3.2, the parse function must be able to parse subexpressions out of the supplied literal body. For example, consider the code snippet in Figure 2.7, expressed instead using the `$rx` TSM:

```
val ssn = $rx /\d\d\d-\d\d-\d\d\d\d/
fun example_rx_tsm(name: string) => $rx /@name: %ssn/
```

The subexpressions `name` and `ssn` on the second line appear directly in the body of the literal form, so we call them *spliced subexpressions* (and color them black when typesetting them in this document). When the parse function determines that a subsequence of the literal body should be treated as a spliced subexpression (here, by recognizing the characters `@` or `%` followed by a variable or parenthesized expression), it can refer to it within the candidate expansion it generates using the `Spliced` constructor of the `CEExp` type shown in Figure 3.3. The `Spliced` constructor requires a value of type `IndexRange` because spliced subexpressions are referred to indirectly by their position within the literal body. This prevents TSMs from "forging" a spliced subexpression (i.e. claiming that an expression is a spliced subexpression, even though it does not appear in the body of the literal form). Expressions can also contain types, so one can also mark spliced types in an analagous manner using the `Spliced` constructor of the `CETyp` type.

The candidate expansion generated by `$rx` for the body of `example_rx_tsm`, if written in a hypothetical concrete syntax for candidate expansions where references to spliced subexpressions are written `spliced<startIdx, endIndex>`, is:

```
Seq(Str(spliced<1, 4>), Seq(Str ": ", spliced<8, 10>))
```

Here, `spliced<1, 4>` refers to the subexpression `name` by position and `spliced<8, 10>` refers to the subexpression `ssn` by position.

3.1.4 Typing

The language *validates* candidate expansions before a final expansion is generated. One aspect of candidate expansion validation is checking the candidate expansion against the type annotation specified by the TSM, e.g. the type `Rx` in the example above. This maintains a *type discipline*: if a programmer sees a TSM being applied when examining a well-typed program, they need only look up the TSM's type annotation to determine the type of the generated expansion. Determining the type does not require examining the expansion directly.

3.1.5 Hygiene

The spliced subexpressions that the candidate expansion refers to (by their position within the literal body, cf. above) must be parsed, typed and expanded during the candidate expansion validation process (otherwise, the language would not be able to check the type of the candidate expansion). To maintain a useful *binding discipline*, i.e. to allow programmers to reason also about variable binding without examining expansions directly, the validation process maintains two additional properties related to spliced subexpressions: **context independent expansion** and **expansion independent splicing**. These are collectively referred to as the *hygiene properties* (because they are conceptually related to the concept of hygiene in term rewriting macro systems, cf. Sec. 2.4.5.)

Context Independent Expansion   Programmers expect to be able to choose variable and symbol names freely, i.e. without needing to satisfy "hidden assumptions" made by the TSMs that are applied in scope of a binding. For this reason, context-dependent candidate expansions, i.e. those with free variables or symbols, are deemed invalid (even at application sites where those variables happen to be bound). An example of a TSM that generates context-dependent candidate expansions is shown below:

```
syntax $bad1 at Rx {
  static fn(body : Body) : ParseResultExp => Success (Var 'x')
}
```

The candidate expansion this TSM generates would be well-typed only when there is an assumption `x : Rx` in the application site typing context. This "hidden assumption" makes reasoning about binding and renaming especially difficult, so this candidate expansion is deemed invalid (even when `$bad1` is applied in a context where `x` happens to be bound).

Of course, this prohibition does not extend into the spliced subexpressions referred to in a candidate expansion because spliced subexpressions are authored by the TSM client and appear at the application site, and so can justifiably refer to application site bindings. We saw examples of spliced subexpressions that referred to variables bound at the application site in Sec. 3.1.3. Because candidate expansions refer to spliced subexpressions indirectly, checking this property is straightforward – we only allow access to the application site typing context when typing spliced subexpressions. In the next section, we will formalize this intuition.

In the examples in Sec. 3.1.1 and Sec. 3.1.3, the expansion used constructors associated with the `Rx` type, e.g. `Seq` and `Str`. This might appear to violate our prohibition on context-dependent expansions. This is not the case only because in VerseML, constructor labels are not variables or scoped symbols. Syntactically, they must begin with a capital letter (like Haskell's datatype constructors). Different labeled sum types can use common constructor labels without conflict because the type the term is being checked against – e.g. `Rx`, due to the type ascription on `$rx` – determines which type of value will be constructed. For dialects of ML where datatype definitions do introduce new

variables or scoped symbols, we need parameterized TSMs. We will return to this topic in Chapter 5.

Expansion Independent Splicing Spliced subexpressions, as just described, must be given access to application site bindings. The *expansion independent splicing* property ensures that spliced subexpressions have access to *only* those bindings, i.e. a TSM cannot introduce new bindings into spliced subexpressions. For example, consider the following hypothetical candidate expansion (written concretely as above):

```
fn(x : Rx) => spliced<0, 4>
```

The variable x is not available when typing the indicated spliced subexpression, nor can it shadow any bindings of x that might appear at the application site.

For TSM providers, the benefit of this property is that they can choose the names of variables used internally within expansions freely, without worrying about whether they might shadow those that a client might have defined at the application site.

TSM clients can, in turn, determine exactly which bindings are available in a spliced subexpression without examining the expansion it appears within. In other words, there can be no "hidden variables".

The trade-off is that this prevents library providers from defining alternative binding forms. For example, Haskell's derived form for monadic commands (i.e. **do**-notation) supports binding the result of executing a command to a variable that is then available in the subsequent commands in a command sequence. In VerseML, this cannot be expressed in the same way. We will show an alternative formulation of Haskell's syntax for monadic commands that uses VerseML's anonymous function syntax to bind variables in Sec. 8.1.1. We will discuss mechanisms that would allow us to relax this restriction while retaining client control over variable names as future work in Sec. 8.6.6.

3.1.6   Final Expansion

If validation succeeds, the language generates the *final expansion* from the candidate expansion by replacing references to spliced subexpressions with their final expansions. The final expansion of the body of `example_rx_tsm` is:

```
Seq(Str(name), Seq(Str ": ", ssn))
```

3.1.7   Scoping

A benefit of specifying TSMs as a language primitive, rather than relying on extralinguistic mechanisms to manipulate the concrete syntax of our language directly, is that TSMs follow standard scoping rules.

For example, we can define a TSM that is visible only to a single expression like this:

```
let x = (syntax $rx at Rx { (* ... *) } in
            (* $rx is in scope here *)
        end)
in (* $rx is no longer in scope *) end
```

If the **in** clause is omitted, the scope of the TSM extends to the end of the current block. We will consider the question of how TSM definitions can be exported from compilation units in Sec. 8.3.1.

3.1.8 Comparison to ML+Rx

Let us compare the VerseML TSM $rx to ML+Rx, the hypothetical syntactic dialect of ML with support for derived forms for regular expressions described in Sec. 2.3.2.

Both ML+Rx and $rx give programmers the ability to use the same standard syntax for constructing regexes, including syntax for splicing in other strings and regexes. In VerseML, however, we incur the additional syntactic cost of explicitly applying the $rx TSM each time we wish to use regex syntax. This cost does not grow with the size of the regex, so it would only be significant in programs that involve a large number of small regexes (which do, of course, exist). In Chapter 7 we will consider a design where even this syntactic cost can be eliminated in certain situations.

The benefit of this approach is that we can easily define other TSMs to use alongside the $rx TSM without needing to consider the possibility of syntactic conflict. Furthermore, programmers can rely on the typing discipline and the hygienic binding discipline described above to reason about programs, including those that contain unfamiliar forms. Put pithily, VerseML helps programmers avoid "conflict and confusion".

3.2 miniVerse$_{UE}$

To make the intuitions developed in the previous section mathematically precise, we will now introduce a reduced language called miniVerse$_{UE}$ with support for ueTSMs. miniVerse$_{UE}$ consists of an *inner core* and an *outer surface*.

3.2.1 Syntax of the Inner Core

The *inner core of* miniVerse$_{UE}$ consists of *types*, τ, and *expanded expressions*, e. The syntax of the inner core is specified by the syntax chart in Figure 3.4. The inner core forms a pure language with support for partial functions, quantification over types, recursive types, labeled product types and labeled sum types. The reader is directed to *PFPL* [29] (or another text on type systems, e.g. *TAPL* [53]) for a detailed introductory account of these (or very similar) constructs. We will tersely define the statics of the inner core, and outline the structural dynamics, in the next two subsections, respectively.

3.2.2 Statics of the Inner Core

The *statics of the inner core* is defined by hypothetical judgements of the following form:

Sort			Operational Form	Stylized Form	Description
Typ	τ	$::=$	t	t	variable
			$\texttt{parr}(\tau;\tau)$	$\tau \rightharpoonup \tau$	partial function
			$\texttt{all}(t.\tau)$	$\forall t.\tau$	polymorphic
			$\texttt{rec}(t.\tau)$	$\mu t.\tau$	recursive
			$\texttt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L})$	$\langle\{i \hookrightarrow \tau_i\}_{i \in L}\rangle$	labeled product
			$\texttt{sum}[L](\{i \hookrightarrow \tau_i\}_{i \in L})$	$[\{i \hookrightarrow \tau_i\}_{i \in L}]$	labeled sum
Exp	e	$::=$	x	x	variable
			$\texttt{lam}\{\tau\}(x.e)$	$\lambda x{:}\tau.e$	abstraction
			$\texttt{ap}(e;e)$	$e(e)$	application
			$\texttt{tlam}(t.e)$	$\Lambda t.e$	type abstraction
			$\texttt{tap}\{\tau\}(e)$	$e[\tau]$	type application
			$\texttt{fold}\{t.\tau\}(e)$	$\texttt{fold}(e)$	fold
			$\texttt{unfold}(e)$	$\texttt{unfold}(e)$	unfold
			$\texttt{tpl}[L](\{i \hookrightarrow e_i\}_{i \in L})$	$\langle\{i \hookrightarrow e_i\}_{i \in L}\rangle$	labeled tuple
			$\texttt{pr}[\ell](e)$	$e \cdot \ell$	projection
			$\texttt{in}[L][\ell]\{\{i \hookrightarrow \tau_i\}_{i \in L}\}(e)$	$\texttt{inj}\ \ell \cdot e$	injection
			$\texttt{case}[L]\{\tau\}(e;\{i \hookrightarrow x_i.e_i\}_{i \in L})$	$\texttt{case}\ e\ \{i \hookrightarrow x_i.e_i\}_{i \in L}$	case analysis

Figure 3.4: Abstract syntax of types and expanded expressions, which form the *inner core of* miniVerse$_{\textbf{UE}}$. Metavariables x range over variables, t over type variables, ℓ over labels and L over finite sets of labels. We adopt *PFPL*'s conventions for operational forms, i.e. the names of operators and indexed families of operators are written in `typewriter font`, indexed families of operators specify non-symbolic indices within [mathematical braces] and symbolic indices within [textual braces], and term arguments are grouped arbitrarily (roughly, by "phase") using {textual curly braces} and (textual rounded braces) [29]. We write $\{i \hookrightarrow \tau_i\}_{i \in L}$ for a sequence of arguments τ_i, one for each $i \in L$, and similarly for arguments of other valences. Operations parameterized by label sets, e.g. $\texttt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L})$, are identified up to mutual reordering of the label set and the corresponding argument sequence. When we use the stylized forms, we assume that the reader can infer suppressed indices and arguments from the surrounding context. Types and expanded expressions are identified up to α-equivalence.

Judgement Form	Description
$\Delta \vdash \tau$ type	τ is a well-formed type assuming Δ
$\Delta\,\Gamma \vdash e : \tau$	e is assigned type τ assuming Δ and Γ

Type formation contexts, Δ, are finite sets of hypotheses of the form t type. Empty finite sets are written $\varnothing$, or omitted entirely within judgements, and non-empty finite sets are written as comma-separated finite sequences identified up to exchange and contraction. We write Δ, t type, when t type $\notin \Delta$, for Δ extended with the hypothesis t type.

The *type formation judgement*, $\Delta \vdash \tau$ type, is inductively defined by the following rules:

$$\frac{}{\Delta, t \text{ type} \vdash t \text{ type}} \tag{3.1a}$$

$$\frac{\Delta \vdash \tau_1 \text{ type} \qquad \Delta \vdash \tau_2 \text{ type}}{\Delta \vdash \texttt{parr}(\tau_1; \tau_2) \text{ type}} \tag{3.1b}$$

$$\frac{\Delta, t \text{ type} \vdash \tau \text{ type}}{\Delta \vdash \texttt{all}(t.\tau) \text{ type}} \tag{3.1c}$$

$$\frac{\Delta, t \text{ type} \vdash \tau \text{ type}}{\Delta \vdash \texttt{rec}(t.\tau) \text{ type}} \tag{3.1d}$$

$$\frac{\{\Delta \vdash \tau_i \text{ type}\}_{i \in L}}{\Delta \vdash \texttt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L}) \text{ type}} \tag{3.1e}$$

$$\frac{\{\Delta \vdash \tau_i \text{ type}\}_{i \in L}}{\Delta \vdash \texttt{sum}[L](\{i \hookrightarrow \tau_i\}_{i \in L}) \text{ type}} \tag{3.1f}$$

Premises of the form $\{J_i\}_{i \in L}$ mean that for each $i \in L$, the judgement J_i must hold.

Typing contexts, Γ, are finite functions that map each variable $x \in \text{dom}(\Gamma)$, to the hypothesis $x : \tau$, for some τ. Empty typing contexts are written $\varnothing$, or omitted entirely within judgements, and non-empty typing contexts are written as finite sequences of hypotheses identified up to exchange (we do not separately write down the finite set $\text{dom}(\Gamma)$ because it can be determined from the listed hypotheses). We write $\Gamma, x : \tau$, when $x \notin \text{dom}(\Gamma)$, for the extension of Γ with a mapping from x to $x : \tau$, and $\Gamma \cup \Gamma'$ when $\text{dom}(\Gamma) \cap \text{dom}(\Gamma') = \varnothing$ for the typing context mapping each $x \in \text{dom}(\Gamma) \cup \text{dom}(\Gamma')$ to $x : \tau$ if $x : \tau \in \Gamma$ or $x : \tau \in \Gamma'$. We write $\Delta \vdash \Gamma$ ctx if every type in Γ is well-formed relative to Δ.

Definition 3.1 (Typing Context Formation). $\Delta \vdash \Gamma$ ctx *iff for each hypothesis* $x : \tau \in \Gamma$, *we have* $\Delta \vdash \tau$ type.

The typing judgement, $\Delta\,\Gamma \vdash e : \tau$, assigns types to expressions. It is inductively defined by the following rules:

$$\frac{}{\Delta\,\Gamma, x : \tau \vdash x : \tau} \tag{3.2a}$$

$$\frac{\Delta \vdash \tau \text{ type} \qquad \Delta\,\Gamma, x : \tau \vdash e : \tau'}{\Delta\,\Gamma \vdash \texttt{lam}\{\tau\}(x.e) : \texttt{parr}(\tau; \tau')} \tag{3.2b}$$

$$\frac{\Delta\,\Gamma \vdash e_1 : \texttt{parr}(\tau; \tau') \qquad \Delta\,\Gamma \vdash e_2 : \tau}{\Delta\,\Gamma \vdash \texttt{ap}(e_1; e_2) : \tau'} \tag{3.2c}$$

$$\frac{\Delta, t \text{ type } \Gamma \vdash e : \tau}{\Delta\,\Gamma \vdash \texttt{tlam}(t.e) : \texttt{all}(t.\tau)} \tag{3.2d}$$

$$\frac{\Delta\,\Gamma \vdash e : \texttt{all}(t.\tau) \qquad \Delta \vdash \tau' \text{ type}}{\Delta\,\Gamma \vdash \texttt{tap}\{\tau'\}(e) : [\tau'/t]\tau} \tag{3.2e}$$

$$\frac{\Delta, t \text{ type} \vdash \tau \text{ type} \qquad \Delta\,\Gamma \vdash e : [\texttt{rec}(t.\tau)/t]\tau}{\Delta\,\Gamma \vdash \texttt{fold}\{t.\tau\}(e) : \texttt{rec}(t.\tau)} \tag{3.2f}$$

$$\frac{\Delta\,\Gamma \vdash e : \texttt{rec}(t.\tau)}{\Delta\,\Gamma \vdash \texttt{unfold}(e) : [\texttt{rec}(t.\tau)/t]\tau} \tag{3.2g}$$

$$\frac{\{\Delta\,\Gamma \vdash e_i : \tau_i\}_{i \in L}}{\Delta\,\Gamma \vdash \texttt{tpl}[L](\{i \hookrightarrow e_i\}_{i \in L}) : \texttt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L})} \tag{3.2h}$$

$$\frac{\Delta\,\Gamma \vdash e : \texttt{prod}[L, \ell](\{i \hookrightarrow \tau_i\}_{i \in L}; \ell \hookrightarrow \tau)}{\Delta\,\Gamma \vdash \texttt{pr}[\ell](e) : \tau} \tag{3.2i}$$

$$\frac{\{\Delta \vdash \tau_i \text{ type}\}_{i \in L} \qquad \Delta \vdash \tau \text{ type} \qquad \Delta\,\Gamma \vdash e : \tau}{\Delta\,\Gamma \vdash \texttt{in}[L, \ell][\ell]\{\{i \hookrightarrow \tau_i\}_{i \in L}; \ell \hookrightarrow \tau\}(e) : \texttt{sum}[L, \ell](\{i \hookrightarrow \tau_i\}_{i \in L}; \ell \hookrightarrow \tau)} \tag{3.2j}$$

$$\frac{\Delta\,\Gamma \vdash e : \texttt{sum}[L](\{i \hookrightarrow \tau_i\}_{i \in L}) \qquad \Delta \vdash \tau \text{ type} \qquad \{\Delta\,\Gamma, x_i : \tau_i \vdash e_i : \tau\}_{i \in L}}{\Delta\,\Gamma \vdash \texttt{case}[L]\{\tau\}(e; \{i \hookrightarrow x_i.e_i\}_{i \in L}) : \tau} \tag{3.2k}$$

Rules (3.1) and (3.2) are syntax-directed, so we assume an inversion lemma for each rule as needed without stating it separately. The following standard lemmas also hold.

The Weakening Lemma establishes that extending a context with unnecessary hypotheses preserves well-formedness and typing.

Lemma 3.2 (Weakening). *All of the following hold:*

1. *If $\Delta \vdash \tau$ type then Δ, t type $\vdash \tau$ type.*
2. *If $\Delta\,\Gamma \vdash e : \tau$ then Δ, t type $\Gamma \vdash e : \tau$.*
3. *If $\Delta\,\Gamma \vdash e : \tau$ and $\Delta \vdash \tau'$ type then $\Delta\,\Gamma, x : \tau' \vdash e : \tau$.*

Proof Sketch. For each part, by rule induction on the assumption. □

We assume that renaming of bound variables, α-equivalence and substitution are defined as in *PFPL* [29]. The Substitution Lemma establishes that substitution of a well-formed type for a type variable, or an expanded expression of the appropriate type for an expanded expression variable, preserves well-formedness and typing.

Lemma 3.3 (Substitution). *All of the following hold:*

1. *If Δ, t type $\vdash \tau$ type and $\Delta \vdash \tau'$ type then $\Delta \vdash [\tau'/t]\tau$ type.*
2. *If Δ, t type $\Gamma \vdash e : \tau$ and $\Delta \vdash \tau'$ type then $\Delta\,[\tau'/t]\Gamma \vdash [\tau'/t]e : [\tau'/t]\tau$.*

3. *If $\Delta\,\Gamma, x : \tau' \vdash e : \tau$ and $\Delta\,\Gamma \vdash e' : \tau'$ then $\Delta\,\Gamma \vdash [e'/x]e : \tau$.*

Proof Sketch. For each part, by rule induction on the first assumption. $\qquad\square$

The Decomposition Lemma is the converse of the Substitution Lemma.

Lemma 3.4 (Decomposition). *All of the following hold:*

1. *If $\Delta \vdash [\tau'/t]\tau$ type and $\Delta \vdash \tau'$ type then Δ, t type $\vdash \tau$ type.*
2. *If $\Delta\,[\tau'/t]\Gamma \vdash [\tau'/t]e : [\tau'/t]\tau$ and $\Delta \vdash \tau'$ type then Δ, t type $\Gamma \vdash e : \tau$.*
3. *If $\Delta\,\Gamma \vdash [e'/x]e : \tau$ and $\Delta\,\Gamma \vdash e' : \tau'$ then $\Delta\,\Gamma, x : \tau' \vdash e : \tau$.*

Proof Sketch.

1. By rule induction over Rules (3.1) and case analysis on the definition of substitution. In all cases, the derivation of $\Delta \vdash [\tau'/t]\tau$ type does not depend on the form of τ'.

2. By rule induction over Rules (3.2) and case analysis on the definition of substitution. In all cases, the derivation of $\Delta\,[\tau'/t]\Gamma \vdash [\tau'/t]e : [\tau'/t]\tau$ does not depend on the form of τ'.

3. By rule induction over Rules (3.2) and case analysis on the definition of substitution. In all cases, the derivation of $\Delta\,\Gamma \vdash [e'/x]e : \tau$ does not depend on the form of e'.

$\qquad\square$

The Regularity Lemma establishes that the type assigned to an expanded expression under a well-formed typing context is always well-formed.

Lemma 3.5 (Regularity). *If $\Delta\,\Gamma \vdash e : \tau$ and $\Delta \vdash \Gamma$ ctx then $\Delta \vdash \tau$ type.*

Proof Sketch. By rule induction over Rules (3.2) and application of Definition 3.1 and Lemma 3.3. $\qquad\square$

3.2.3 Structural Dynamics

The *structural dynamics of* miniVerse$_{\mathbf{UE}}$ is specified as a transition system by judgements of the following form:

Judgement Form	Description
$e \mapsto e'$	e transitions to e'
e val	e is a value

We also define auxiliary judgements for *iterated transition*, $e \mapsto^* e'$, and *evaluation*, $e \Downarrow e'$.

Definition 3.6 (Iterated Transition). *$e \mapsto^* e'$ is the reflexive, transitive closure of $e \mapsto e'$.*

Definition 3.7 (Evaluation). *$e \Downarrow e'$ iff $e \mapsto^* e'$ and e' val.*

Our subsequent developments do not require making reference to particular rules in the structural dynamics (because TSMs operate statically), so we do not reproduce the rules here. Instead, it suffices to state the following conditions.

The Canonical Forms condition characterizes well-typed values. Satisfying this condition requires an *eager* (i.e. *by-value*) formulation of the dynamics.

Condition 3.8 (Canonical Forms). *If $\vdash e : \tau$ and e val then:*

1. *If $\tau = \mathtt{parr}(\tau_1; \tau_2)$ then $e = \mathtt{lam}\{\tau_1\}(x.e')$ and $x : \tau_1 \vdash e' : \tau_2$.*

2. *If* $\tau = \mathtt{all}(t.\tau')$ *then* $e = \mathtt{tlam}(t.e')$ *and* t type $\vdash e' : \tau'$.

3. *If* $\tau = \mathtt{rec}(t.\tau')$ *then* $e = \mathtt{fold\{t.\tau'\}}(e')$ *and* $\vdash e' : [\mathtt{rec}(t.\tau')/t]\tau'$ *and* e' val.

4. *If* $\tau = \mathtt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L})$ *then* $e = \mathtt{tpl}[L](\{i \hookrightarrow e_i\}_{i \in L})$ *and* $\vdash e_i : \tau_i$ *and* e_i val *for each* $i \in L$.

5. *If* $\tau = \mathtt{sum}[L](\{i \hookrightarrow \tau_i\}_{i \in L})$ *then for some label set L' and label ℓ and type τ_ℓ, we have that* $L = L', \ell$ *and* $\tau = \mathtt{sum}[L', \ell](\{i \hookrightarrow \tau_i\}_{i \in L'}; \ell \hookrightarrow \tau_\ell)$ *and* $e = \mathtt{in}[L', \ell][\ell]\{\{i \hookrightarrow \tau_i\}_{i \in L'}; \ell \hookrightarrow \tau_\ell\}(e')$ *and* $\vdash e' : \tau_\ell$ *and* e' val.

The Preservation condition ensures that evaluation preserves typing.

Condition 3.9 (Preservation). *If* $\vdash e : \tau$ *and* $e \mapsto^* e'$ *then* $\vdash e' : \tau$.

The Progress condition ensures that evaluation of a well-typed expanded expression cannot "get stuck".

Condition 3.10 (Progress). *If* $\vdash e : \tau$ *then either* e val *or there exists an* e' *such that* $e \mapsto e'$.

Together, these two conditions constitute the Type Safety Condition.

3.2.4 Syntax of the Outer Surface

A miniVerse$_{\mathbf{UE}}$ program ultimately evaluates as an expanded expression. However, the programmer does not write the expanded expression directly. Instead, the programmer writes a textual sequence, b, consisting of characters in some suitable alphabet (e.g. in practice, ASCII or Unicode), which is parsed by some partial metafunction parseUExp(b) to produce an *unexpanded expression, $\hat{e}$*. Unexpanded expressions can contain *unexpanded types, $\hat{\tau}$*, so we also need a partial metafunction parseUTyp(b). The abstract syntax of unexpanded types and expressions, which form the *outer surface* of miniVerse$_{\mathbf{UE}}$, is defined in Figure 3.5. The full definition of the textual syntax of miniVerse$_{\mathbf{UE}}$, which parseUExp(b) and parseUTyp(b) implement, is not important for our purposes, so we simply give the following condition, which states that there is some way to textually represent every unexpanded type and expression.

Condition 3.11 (Textual Representability). *Both of the following must hold:*

1. *For each $\hat{\tau}$, there exists b such that* parseUTyp(b) $= \hat{\tau}$.

2. *For each $\hat{e}$, there exists b such that* parseUExp(b) $= \hat{e}$.

Unexpanded types and expressions are given meaning by expansion to types and expanded expressions, respectively, according to the *typed expansion judgements*, which are defined in the next subsection.

Unexpanded types and expressions bind *type sigils, $\hat{t}$, expression sigils, $\hat{x}$,* and *TSM names, $\hat{a}$*. Sigils are given meaning by expansion to variables during typed expansion. We **cannot** adopt the usual definitions of α-renaming of identifiers, because unexpanded types and expressions are still in a "partially parsed" state – the literal bodies, b, within an unexpanded expression might contain spliced subterms that are "surfaced" by a TSM only during typed expansion, as we will detail below.

Each inner core form (defined in Figure 3.4) maps onto an outer surface form. We refer to these as the *shared forms*. In particular:

- Each type variable, t, maps onto a unique type sigil, written $\widehat{t}$ (pronounced "sigil of t"). Notice the distinction between $\hat{t}$, which is a metavariable ranging over type

Sort	Operational Form	Stylized Form	Description
UTyp $\hat{\tau}$ ::=	$\hat{t}$	$\hat{t}$	sigil
	$\mathtt{uarr}(\hat{\tau};\hat{\tau})$	$\hat{\tau} \rightharpoonup \hat{\tau}$	partial function
	$\mathtt{uall}(\hat{t}.\hat{\tau})$	$\forall \hat{t}.\hat{\tau}$	polymorphic
	$\mathtt{urec}(\hat{t}.\hat{\tau})$	$\mu\hat{t}.\hat{\tau}$	recursive
	$\mathtt{uprod}[L](\{i \hookrightarrow \hat{\tau}_i\}_{i \in L})$	$\langle\{i \hookrightarrow \hat{\tau}_i\}_{i \in L}\rangle$	labeled product
	$\mathtt{usum}[L](\{i \hookrightarrow \hat{\tau}_i\}_{i \in L})$	$[\{i \hookrightarrow \hat{\tau}_i\}_{i \in L}]$	labeled sum
UExp $\hat{e}$ ::=	$\hat{x}$	$\hat{x}$	sigil
	$\mathtt{ulam}\{\hat{\tau}\}(\hat{x}.\hat{e})$	$\lambda\hat{x}{:}\hat{\tau}.\hat{e}$	abstraction
	$\mathtt{uap}(\hat{e};\hat{e})$	$\hat{e}(\hat{e})$	application
	$\mathtt{utlam}(\hat{t}.\hat{e})$	$\Lambda\hat{t}.\hat{e}$	type abstraction
	$\mathtt{utap}\{\hat{\tau}\}(\hat{e})$	$\hat{e}[\hat{\tau}]$	type application
	$\mathtt{ufold}\{\hat{t}.\hat{\tau}\}(\hat{e})$	$\mathtt{fold}(\hat{e})$	fold
	$\mathtt{uunfold}(\hat{e})$	$\mathtt{unfold}(\hat{e})$	unfold
	$\mathtt{utpl}[L](\{i \hookrightarrow \hat{e}_i\}_{i \in L})$	$\langle\{i \hookrightarrow \hat{e}_i\}_{i \in L}\rangle$	labeled tuple
	$\mathtt{upr}[\ell](\hat{e})$	$\hat{e} \cdot \ell$	projection
	$\mathtt{uin}[L][\ell]\{\{i \hookrightarrow \hat{\tau}_i\}_{i \in L}\}(\hat{e})$	$\mathtt{inj}\ \ell \cdot \hat{e}$	injection
	$\mathtt{ucase}[L]\{\hat{\tau}\}(\hat{e};\{i \hookrightarrow \hat{x}_i.\hat{e}_i\}_{i \in L})$	$\mathtt{case}\ \hat{e}\ \{i \hookrightarrow \hat{x}_i.\hat{e}_i\}_{i \in L}$	case analysis
	$\mathtt{usyntaxue}\{e\}\{\hat{\tau}\}(\hat{a}.\hat{e})$	$\mathtt{syntax}\ \hat{a}\ \mathtt{at}\ \hat{\tau}\ \{e\}\ \mathtt{in}\ \hat{e}$	ueTSM definition
	$\mathtt{uapuetsm}[b][\hat{a}]$	$\hat{a}\ /b/$	ueTSM application

Figure 3.5: Abstract syntax of unexpanded types and expressions in miniVerse$_{\mathsf{UE}}$. Metavariable $\hat{t}$ ranges over type sigils, $\hat{x}$ ranges over expression sigils, $\hat{a}$ over TSM names and b over textual sequences, which, when they appear in an unexpanded expression, are called literal bodies. Literal bodies might contain spliced subterms that are only "surfaced" during typed expansion, so renaming of bound identifiers and substitution are not defined over unexpanded types and expressions.

sigils, and $\widehat{t}$, which is a metafunction, written in stylized form, applied to a type variable to produce a type sigil.

- Each type form, τ, maps onto an unexpanded type form, $\mathcal{U}(\tau)$, as follows:

$$\mathcal{U}(t) = \widehat{t}$$
$$\mathcal{U}(\mathtt{parr}(\tau_1; \tau_2)) = \mathtt{uparr}(\mathcal{U}(\tau_1); \mathcal{U}(\tau_2))$$
$$\mathcal{U}(\mathtt{all}(t.\tau)) = \mathtt{uall}(\widehat{t}.\mathcal{U}(\tau))$$
$$\mathcal{U}(\mathtt{rec}(t.\tau)) = \mathtt{urec}(\widehat{t}.\mathcal{U}(\tau))$$
$$\mathcal{U}(\mathtt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L})) = \mathtt{uprod}[L](\{i \hookrightarrow \mathcal{U}(\tau_i)\}_{i \in L})$$
$$\mathcal{U}(\mathtt{sum}[L](\{i \hookrightarrow \tau_i\}_{i \in L})) = \mathtt{usum}[L](\{i \hookrightarrow \mathcal{U}(\tau_i)\}_{i \in L})$$

- Each expression variable, x, maps onto a unique expression sigil written $\widehat{x}$. Again, notice the distinction between $\hat{x}$ and $\widehat{x}$.

- Each expanded expression form, e, maps onto an unexpanded expression form, $\mathcal{U}(e)$, as follows:

$$\mathcal{U}(x) = \widehat{x}$$
$$\mathcal{U}(\mathtt{lam}\{\tau\}(x.e)) = \mathtt{ulam}\{\mathcal{U}(\tau)\}(\widehat{x}.\mathcal{U}(e))$$
$$\mathcal{U}(\mathtt{ap}(e_1; e_2)) = \mathtt{uap}(\mathcal{U}(e_1); \mathcal{U}(e_2))$$
$$\mathcal{U}(\mathtt{tlam}(t.e)) = \mathtt{utlam}(\widehat{t}.\mathcal{U}(e))$$
$$\mathcal{U}(\mathtt{tap}\{\tau\}(e)) = \mathtt{utap}\{\mathcal{U}(\tau)\}(\mathcal{U}(e))$$
$$\mathcal{U}(\mathtt{fold}\{t.\tau\}(e)) = \mathtt{ufold}\{\widehat{t}.\mathcal{U}(\tau)\}(\mathcal{U}(e))$$
$$\mathcal{U}(\mathtt{unfold}(e)) = \mathtt{uunfold}(\mathcal{U}(e))$$
$$\mathcal{U}(\mathtt{tpl}[L](\{i \hookrightarrow e_i\}_{i \in L})) = \mathtt{utpl}[L](\{i \hookrightarrow \mathcal{U}(e_i)\}_{i \in L})$$
$$\mathcal{U}(\mathtt{in}[L][\ell]\{\{i \hookrightarrow \tau_i\}_{i \in L}\}(e)) = \mathtt{uin}[L][\ell]\{\{i \hookrightarrow \mathcal{U}(\tau_i)\}_{i \in L}\}(\mathcal{U}(e))$$
$$\mathcal{U}(\mathtt{case}[L]\{\tau\}(e; \{i \hookrightarrow x_i.e_i\}_{i \in L})) = \mathtt{ucase}[L]\{\mathcal{U}(\tau)\}(\mathcal{U}(e); \{i \hookrightarrow \widehat{x}_i.\mathcal{U}(e_i)\}_{i \in L})$$

There are only two unexpanded expression forms, highlighted in gray in Figure 3.5, that do not correspond to expanded expression forms – the ueTSM definition form and the ueTSM application form.

3.2.5 Typed Expansion

Unexpanded expressions, and the unexpanded types therein, are checked and expanded simultaneously according to the *typed expansion judgements*:

Judgement Form	Description
$\hat{\Delta} \vdash \hat{\tau} \rightsquigarrow \tau$ type	$\hat{\tau}$ is well-formed and has expansion τ assuming $\hat{\Delta}$
$\hat{\Delta}\, \hat{\Gamma} \vdash_{\hat{\Psi}} \hat{e} \rightsquigarrow e : \tau$	$\hat{e}$ has expansion e and type τ under ueTSM context $\hat{\Psi}$ assuming $\hat{\Delta}$ and $\hat{\Gamma}$

Type Expansion

The *type expansion judgement*, $\hat{\Delta} \vdash \hat{\tau} \rightsquigarrow \tau$ type, is inductively defined by the following rules.

$$\frac{}{\hat{\Delta}, \hat{t} \rightsquigarrow t \text{ type} \vdash \hat{t} \rightsquigarrow t \text{ type}} \tag{3.3a}$$

$$\frac{\hat{\Delta} \vdash \hat{\tau}_1 \rightsquigarrow \tau_1 \text{ type} \qquad \hat{\Delta} \vdash \hat{\tau}_2 \rightsquigarrow \tau_2 \text{ type}}{\hat{\Delta} \vdash \texttt{uparr}(\hat{\tau}_1; \hat{\tau}_2) \rightsquigarrow \texttt{parr}(\tau_1; \tau_2) \text{ type}} \tag{3.3b}$$

$$\frac{\hat{\Delta}, \hat{t} \rightsquigarrow t \text{ type} \vdash \hat{\tau} \rightsquigarrow \tau \text{ type}}{\hat{\Delta} \vdash \texttt{uall}(\hat{t}.\hat{\tau}) \rightsquigarrow \texttt{all}(t.\tau) \text{ type}} \tag{3.3c}$$

$$\frac{\hat{\Delta}, \hat{t} \rightsquigarrow t \text{ type} \vdash \hat{\tau} \rightsquigarrow \tau \text{ type}}{\hat{\Delta} \vdash \texttt{urec}(\hat{t}.\hat{\tau}) \rightsquigarrow \texttt{rec}(t.\tau) \text{ type}} \tag{3.3d}$$

$$\frac{\{\hat{\Delta} \vdash \hat{\tau}_i \rightsquigarrow \tau_i \text{ type}\}_{i \in L}}{\hat{\Delta} \vdash \texttt{uprod}[L](\{i \hookrightarrow \hat{\tau}_i\}_{i \in L}) \rightsquigarrow \texttt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L}) \text{ type}} \tag{3.3e}$$

$$\frac{\{\hat{\Delta} \vdash \hat{\tau}_i \rightsquigarrow \tau_i \text{ type}\}_{i \in L}}{\hat{\Delta} \vdash \texttt{usum}[L](\{i \hookrightarrow \hat{\tau}_i\}_{i \in L}) \rightsquigarrow \texttt{sum}[L](\{i \hookrightarrow \tau_i\}_{i \in L}) \text{ type}} \tag{3.3f}$$

Unexpanded type formation contexts, $\hat{\Delta}$, are of the form $\langle \mathcal{D}; \Delta \rangle$, where $\mathcal{D}$ is a *type sigil expansion context*, and Δ is a type formation context. A type sigil expansion context, $\mathcal{D}$, is a finite function that maps each type sigil $\hat{t} \in \text{dom}(\mathcal{D})$ to the hypothesis $\hat{t} \rightsquigarrow t$, for some type variable t. We write $\mathcal{D} \uplus \hat{t} \rightsquigarrow t$ for the type sigil expansion context that maps $\hat{t}$ to $\hat{t} \rightsquigarrow t$ and defers to $\mathcal{D}$ for all other type sigils (i.e. the previous mapping, if it exists, is updated). We define $\hat{\Delta}, \hat{t} \rightsquigarrow t$ type when $\hat{\Delta} = \langle \mathcal{D}; \Delta \rangle$ as an abbreviation of

$$\langle \mathcal{D} \uplus \hat{t} \rightsquigarrow t; \Delta, t \text{ type} \rangle$$

To understand how type sigil expansion contexts operate, it is instructive to derive an expansion for the unexpanded type $\forall \hat{t}. \forall \hat{t}. \hat{t}$, or in operational form, $\texttt{uall}(\hat{t}.\texttt{uall}(\hat{t}.\hat{t}))$:

$$\cfrac{\cfrac{\cfrac{}{\langle \hat{t} \rightsquigarrow t'; t \text{ type}, t' \text{ type} \rangle \vdash \hat{t} \rightsquigarrow t' \text{ type}} \text{ (3.3a)}}{\langle \hat{t} \rightsquigarrow t; t \text{ type} \rangle \vdash \texttt{uall}(\hat{t}.\hat{t}) \rightsquigarrow \texttt{all}(t'.t') \text{ type}} \text{ (3.3c)}}{\langle \varnothing; \varnothing \rangle \vdash \texttt{uall}(\hat{t}.\texttt{uall}(\hat{t}.\hat{t})) \rightsquigarrow \texttt{all}(t.\texttt{all}(t'.t')) \text{ type}} \text{ (3.3c)}$$

Notice that when a type sigil is bound, a fresh type variable is generated. The type sigil expansion context is extended (when the outermost binding is encountered) or updated (at all inner bindings) and the type formation context is simultaneously extended at each binding (so that typing contexts and ueTSM contexts, discussed below, that contain types that refer to the previous binding remain well-formed). Had we used type variables in the syntax and type formation contexts in the rules above, rather than type sigils and type sigil expansion contexts, derivations for unexpanded types where an

inner binding shadows an outer binding would not exist, because by definition we cannot extend a type formation context with a variable it already mentions nor implicitly α-vary the unexpanded type to sidestep this problem.

These rules validate the following lemmas. The Type Expansion Lemma establishes that the expansion of an unexpanded type is a well-formed type.

Lemma 3.12 (Type Expansion). *If* $\langle \mathcal{D}; \Delta \rangle \vdash \hat{\tau} \rightsquigarrow \tau$ type *then* $\Delta \vdash \tau$ type.

Proof. By rule induction over Rules (3.3). In each case, we apply the IH to or over each premise, then apply the corresponding type formation rule in Rules (3.1). □

The Type Expressibility Lemma establishes that every well-formed type, τ, can be expressed as a well-formed unexpanded type, $\mathcal{U}(\tau)$. This requires defining the meta-function $\mathcal{U}(\Delta)$ which maps Δ onto a an unexpanded type formation context as follows:

$$\mathcal{U}(\varnothing) = \langle \varnothing; \varnothing \rangle$$
$$\mathcal{U}(\Delta, t \text{ type}) = \mathcal{U}(\Delta), \hat{t} \rightsquigarrow t \text{ type}$$

Lemma 3.13 (Type Expressibility). *If* $\Delta \vdash \tau$ type *then* $\mathcal{U}(\Delta) \vdash \mathcal{U}(\tau) \rightsquigarrow \tau$ type.

Proof. By rule induction over Rules (3.1) using the definitions of $\mathcal{U}(\tau)$ and $\mathcal{U}(\Delta)$ above. In each case, we apply the IH to or over each premise, then apply the corresponding type expansion rule in Rules (3.3). □

Typed Expression Expansion

Unexpanded typing contexts, $\hat{\Gamma}$, are of the form $\langle \mathcal{G}; \Gamma \rangle$, where $\mathcal{G}$ is an *expression sigil expansion context*, and Γ is a typing context. An expression sigil expansion context, $\mathcal{G}$, is a finite function that maps each expression sigil $\hat{x} \in \text{dom}(\mathcal{G})$ to the hypothesis $\hat{x} \rightsquigarrow x$, for some expression variable, x. We write $\mathcal{G} \uplus \hat{x} \rightsquigarrow x$ for the expression sigil expansion context that maps $\hat{x}$ to $\hat{x} \rightsquigarrow x$ and defers to $\mathcal{G}$ for all other expression sigils (i.e. the previous mapping, if it exists, is updated). We define $\hat{\Gamma}, \hat{x} \rightsquigarrow x : \tau$ when $\hat{\Gamma} = \langle \mathcal{G}; \Gamma \rangle$ as an abbreviation of

$$\langle \mathcal{G}, \hat{x} \rightsquigarrow x; \Gamma, x : \tau \rangle$$

The *typed expression expansion judgement*, $\hat{\Delta} \, \hat{\Gamma} \vdash_{\hat{\Psi}} \hat{e} \rightsquigarrow e : \tau$, is inductively defined by Rules (3.4) as follows.

Shared Forms   Rules (3.4a) through (3.4k) handle unexpanded expressions of shared form. The first five of these rules are defined below:

$$\frac{}{\hat{\Delta} \, \hat{\Gamma}, \hat{x} \rightsquigarrow x : \tau \vdash_{\hat{\Psi}} \hat{x} \rightsquigarrow x : \tau} \tag{3.4a}$$

$$\frac{\hat{\Delta} \vdash \hat{\tau} \rightsquigarrow \tau \text{ type} \qquad \hat{\Delta} \, \hat{\Gamma}, \hat{x} \rightsquigarrow x : \tau \vdash_{\hat{\Psi}} \hat{e} \rightsquigarrow e : \tau'}{\hat{\Delta} \, \hat{\Gamma} \vdash_{\hat{\Psi}} \texttt{ulam}\{\hat{\tau}\}(\hat{x}.\hat{e}) \rightsquigarrow \texttt{lam}\{\tau\}(x.e) : \texttt{parr}(\tau; \tau')} \tag{3.4b}$$

$$\frac{\hat{\Delta}\ \hat{\Gamma} \vdash_{\hat{\Psi}} \hat{e}_1 \rightsquigarrow e_1 : \mathtt{parr}(\tau;\tau') \qquad \hat{\Delta}\ \hat{\Gamma} \vdash_{\hat{\Psi}} \hat{e}_2 \rightsquigarrow e_2 : \tau}{\hat{\Delta}\ \hat{\Gamma} \vdash_{\hat{\Psi}} \mathtt{uap}(\hat{e}_1;\hat{e}_2) \rightsquigarrow \mathtt{ap}(e_1;e_2) : \tau'} \tag{3.4c}$$

$$\frac{\hat{\Delta}, \hat{t} \rightsquigarrow t\ \mathsf{type}\ \hat{\Gamma} \vdash_{\hat{\Psi}} \hat{e} \rightsquigarrow e : \tau}{\hat{\Delta}\ \hat{\Gamma} \vdash_{\hat{\Psi}} \mathtt{utlam}(\hat{t}.\hat{e}) \rightsquigarrow \mathtt{tlam}(t.e) : \mathtt{all}(t.\tau)} \tag{3.4d}$$

$$\frac{\hat{\Delta}\ \hat{\Gamma} \vdash_{\hat{\Psi}} \hat{e} \rightsquigarrow e : \mathtt{all}(t.\tau) \qquad \hat{\Delta} \vdash \hat{\tau}' \rightsquigarrow \tau'\ \mathsf{type}}{\hat{\Delta}\ \hat{\Gamma} \vdash_{\hat{\Psi}} \mathtt{utap}\{\hat{\tau}'\}(\hat{e}) \rightsquigarrow \mathtt{tap}\{\tau'\}(e) : [\tau'/t]\tau} \tag{3.4e}$$

Observe that, in each of these rules, the unexpanded and expanded expression forms in the conclusion correspond, and the premises correspond to those of the typing rule for the expanded expression form, i.e. Rules (3.2a) through (3.2e), respectively. In particular, each type expansion premise in each rule above corresponds to a type formation premise in the corresponding typing rule, and each typed expression expansion premise in each rule above corresponds to a typing premise in the corresponding typing rule. The type assigned in the conclusion of each rule above is identical to the type assigned in the conclusion of the corresponding typing rule. The ueTSM context, $\hat{\Psi}$, passes opaquely through these rules (we will define ueTSM contexts below). Rules (3.3) were similarly generated by mechanically transforming Rules (3.1).

We can express this scheme more precisely with the following rule transformation. For each rule in Rules (3.1) and Rules (3.2),

$$\frac{J_1 \qquad \cdots \qquad J_k}{J}$$

the corresponding typed expansion rule is

$$\frac{\mathcal{U}(J_1) \qquad \cdots \qquad \mathcal{U}(J_k)}{\mathcal{U}(J)}$$

where

$$\mathcal{U}(\Delta \vdash \tau\ \mathsf{type}) = \mathcal{U}(\Delta) \vdash \mathcal{U}(\tau) \rightsquigarrow \tau\ \mathsf{type}$$
$$\mathcal{U}(\Gamma\ \Delta \vdash e : \tau) = \mathcal{U}(\Gamma)\ \mathcal{U}(\Delta) \vdash_{\hat{\Psi}} \mathcal{U}(e) \rightsquigarrow e : \tau$$
$$\mathcal{U}(\{J_i\}_{i \in L}) = \{\mathcal{U}(J_i)\}_{i \in L}$$

and where:
- $\mathcal{U}(\tau)$ is defined as follows:
    - When τ is of definite form, $\mathcal{U}(\tau)$ is defined as in Sec. 3.2.4.
    - When τ is of indefinite form, $\mathcal{U}(\tau)$ is a uniquely corresponding metavariable of sort UTyp also of indefinite form. For example, in Rule (3.1b), τ_1 and τ_2 are of indefinite form, i.e. they match arbitrary types. The rule transformation simply "hats" them, i.e. $\mathcal{U}(\tau_1) = \hat{\tau}_1$ and $\mathcal{U}(\tau_2) = \hat{\tau}_2$.
- $\mathcal{U}(e)$ is defined as follows

- When e is of definite form, $\mathcal{U}(e)$ is defined as in Sec. 3.2.4.
- When e is of indefinite form, $\mathcal{U}(e)$ is a uniquely corresponding metavariable of sort UExp also of indefinite form. For example, $\mathcal{U}(e_1) = \hat{e}_1$ and $\mathcal{U}(e_2) = \hat{e}_2$.

- $\mathcal{U}(\Delta)$ is defined as follows:
  - When Δ is of definite form, $\mathcal{U}(\Delta)$ is defined as above.
  - When Δ is of indefinite form, $\mathcal{U}(\Delta)$ is a uniquely corresponding metavariable ranging over unexpanded type formation contexts. For example, $\mathcal{U}(\Delta) = \hat{\Delta}$.

- $\mathcal{U}(\Gamma)$ is defined as follows:
  - When Γ is of definite form, $\mathcal{U}(\Gamma)$ produces the corresponding unexpanded typing context as follows:

$$\mathcal{U}(\varnothing) = \langle \varnothing; \varnothing \rangle$$
$$\mathcal{U}(\Gamma, x : \tau) = \mathcal{U}(\Gamma), \hat{x} \rightsquigarrow x : \tau$$

  - When Γ is of indefinite form, $\mathcal{U}(\Gamma)$ is a uniquely corresponding metavariable ranging over unexpanded typing contexts. For example, $\mathcal{U}(\Gamma) = \hat{\Gamma}$.

It is instructive to use this rule transformation to generate Rules (3.3) and Rules (3.4a) through (3.4e) above. We omit the remaining rules, i.e. Rules (3.4f) through (3.4k). By instead defining these rules solely by the rule transformation just described, we avoid having to write down a number of rules that are of limited marginal interest. Moreover, this demonstrates the general technique for generating typed expansion rules for unexpanded types and expressions of shared form, so our exposition is somewhat "robust" to changes to the inner core.

We can now establish the Expressibility Theorem – that each well-typed expanded expression, e, can be expressed as an unexpanded expression, $\hat{e}$, and assigned the same type under the corresponding contexts.

Theorem 3.14 (Expressibility). *If $\Delta \, \Gamma \vdash e : \tau$ then $\mathcal{U}(\Delta) \, \mathcal{U}(\Gamma) \vdash_{\hat{\Psi}} \mathcal{U}(e) \rightsquigarrow e : \tau$.*

Proof. By rule induction over Rules (3.2). The above rule transformation guarantees that this theorem holds by its construction. In particular, in each case, we can apply Lemma 3.13 to or over each type formation premise, the IH to or over each typing premise, then apply the corresponding rule in Rules (3.4). □

ueTSM Definition and Application   The two remaining typed expansion rules, Rules (3.4l) and (3.4m), govern the ueTSM definition and application forms, and are defined in the next two subsections, respectively.

3.2.6   ueTSM Definitions

The stylized ueTSM definition form is

$$\texttt{syntax } \hat{a} \texttt{ at } \hat{\tau} \; \{e_{\text{parse}}\} \texttt{ in } \hat{e}$$

An unexpanded expression of this form defines a ueTSM named $\hat{a}$ with *unexpanded type annotation* $\hat{\tau}$ and *parse function* e_{parse} for use within $\hat{e}$.

The parse function is an expanded expression because parse functions are applied statically (i.e. during typed expansion of $\hat{e}$), as we will discuss when describing ueTSM application below, and evaluation is defined only for closed expanded expressions. This construction simplifies our exposition, though it is not entirely practical because it provides no way for TSM providers to share values between parse functions, nor any way to use TSMs when defining other TSMs. We discuss enriching the language to eliminate these limitations in Sec. 6.2, but it is pedagogically simpler to leave the necessary machinery out of our calculus for now.

Rule (3.4l) defines typed expansion of ueTSM definitions (we use stylized forms for clarity):

$$\frac{\hat{\Delta} \vdash \hat{\tau} \rightsquigarrow \tau\ \text{type} \qquad \emptyset\ \emptyset \vdash e_{\text{parse}} : \texttt{Body} \rightharpoonup \texttt{ParseResultExp} \qquad \hat{\Delta}\ \hat{\Gamma} \vdash_{\hat{\Psi}, \hat{a} \rightsquigarrow a \hookrightarrow \texttt{uetsm}(\tau; e_{\text{parse}})} \hat{e} \rightsquigarrow e : \tau'}{\hat{\Delta}\ \hat{\Gamma} \vdash_{\hat{\Psi}} \texttt{syntax}\ \hat{a}\ \texttt{at}\ \hat{\tau}\ \{e_{\text{parse}}\}\ \texttt{in}\ \hat{e} \rightsquigarrow e : \tau'} \tag{3.4l}$$

The premises of this rule can be understood as follows, in order:

1. The first premise ensures that the unexpanded type annotation is well-formed and expands it to produce the *type annotation*, τ.

2. The second premise checks that the parse function, e_{parse}, is closed and of type

$$\texttt{Body} \rightharpoonup \texttt{ParseResultExp}$$

The type abbreviated $\texttt{Body}$ classifies encodings of literal bodies, b. The mapping from literal bodies to values of type $\texttt{Body}$ is defined by the *body encoding judgement* $b \downarrow e_{\text{body}}$. An inverse mapping is defined by the *body decoding judgement* $e_{\text{body}} \uparrow b$.

Judgement Form	Description
$b \downarrow e$	b has encoding e
$e \uparrow b$	e has decoding b

Rather than defining $\texttt{Body}$ explicitly, and these judgements inductively against that definition (which would be tedious and uninteresting), it suffices to define the following condition, which establishes an isomorphism between literal bodies and values of type $\texttt{Body}$ mediated by the judgements above.

Condition 3.15 (Body Isomorphism)**.** *All of the following must hold:*

(a) *For every literal body b, we have that $b \downarrow e_{body}$ for some e_{body} such that $\vdash e_{body} : \texttt{Body}$ and e_{body} val.*

(b) *If $\vdash e_{body} : \texttt{Body}$ and e_{body} val then $e_{body} \uparrow b$ for some b.*

(c) *If $b \downarrow e_{body}$ then $e_{body} \uparrow b$.*

(d) *If $\vdash e_{body} : \texttt{Body}$ and e_{body} val and $e_{body} \uparrow b$ then $b \downarrow e_{body}$.*

(e) *If $b \downarrow e_{body}$ and $b \downarrow e'_{body}$ then $e_{body} = e'_{body}$.*

(f) *If $\vdash e_{body} : \texttt{Body}$ and e_{body} val and $e_{body} \uparrow b$ and $e_{body} \uparrow b'$ then $b = b'$.*

`ParseResultExp` abbreviates a labeled sum type that distinguishes successful parses from parse errors[1]:

$$\texttt{ParseResultExp} \triangleq [\texttt{Success} \hookrightarrow \texttt{CEExp}, \texttt{ParseError} \hookrightarrow \langle\rangle]$$

The type abbreviated `CEExp` classifies encodings of *candidate expansion expressions* (or *ce-expressions*), $\grave{e}$ (pronounced "grave e"). The syntax of ce-expressions will be described in Sec. 3.2.8. The mapping from ce-expressions to values of type `CEExp` is defined by the *ce-expression encoding judgement*, $\grave{e} \downarrow_{\textsf{CEExp}} e$. An inverse mapping is defined by the *ce-expression decoding judgement*, $e \uparrow_{\textsf{CEExp}} \grave{e}$.

Judgement Form	Description
$\grave{e} \downarrow_{\textsf{CEExp}} e$	$\grave{e}$ has encoding e
$e \uparrow_{\textsf{CEExp}} \grave{e}$	e has decoding $\grave{e}$

Again, rather than picking a particular definition of `CEExp` and defining the judgements above inductively against it, we only state the following condition, which establishes an isomorphism between values of type `CEExp` and ce-expressions.

Condition 3.16 (Candidate Expansion Expression Isomorphism)**.** *All of the following must hold:*

(a) *For every $\grave{e}$, we have $\grave{e} \downarrow_{\textsf{CEExp}} e_{cand}$ for some e_{cand} such that $\vdash e_{cand} : \texttt{CEExp}$ and e_{cand}* val.

(b) *If $\vdash e_{cand} : \texttt{CEExp}$ and e_{cand} val then $e_{cand} \uparrow_{\textsf{CEExp}} \grave{e}$ for some $\grave{e}$.*

(c) *If $\grave{e} \downarrow_{\textsf{CEExp}} e_{cand}$ then $e_{cand} \uparrow_{\textsf{CEExp}} \grave{e}$.*

(d) *If $\vdash e_{cand} : \texttt{CEExp}$ and e_{cand} val and $e_{cand} \uparrow_{\textsf{CEExp}} \grave{e}$ then $\grave{e} \downarrow_{\textsf{CEExp}} e_{cand}$.*

(e) *If $\grave{e} \downarrow_{\textsf{CEExp}} e_{cand}$ and $\grave{e} \downarrow_{\textsf{CEExp}} e'_{cand}$ then $e_{cand} = e'_{cand}$.*

(f) *If $\vdash e_{cand} : \texttt{CEExp}$ and e_{cand} val and $e_{cand} \uparrow_{\textsf{CEExp}} \grave{e}$ and $e_{cand} \uparrow_{\textsf{CEExp}} \grave{e}'$ then $\grave{e} = \grave{e}'$.*

3. The final premise of Rule (3.4l) extends the ueTSM context, $\hat{\Psi}$, with the newly determined ueTSM definition, and proceeds to assign a type, τ', and expansion, e, to $\hat{e}$. The conclusion of Rule (3.4l) assigns this type and expansion to the ueTSM definition as a whole.

ueTSM contexts, $\hat{\Psi}$, are of the form $\langle \mathcal{A}; \Psi \rangle$, where $\mathcal{A}$ is a *TSM naming context* and Ψ is a *ueTSM definition context*.

A *TSM naming context*, $\mathcal{A}$, is a finite function mapping each TSM name $\hat{a} \in \text{dom}(\mathcal{A})$ to the *TSM name-symbol mapping*, $\hat{a} \rightsquigarrow a$, for some *symbol*, a. We write $\mathcal{A} \uplus \hat{a} \rightsquigarrow a$ for the ueTSM naming context that maps $\hat{a}$ to $\hat{a} \rightsquigarrow a$, and defers to $\mathcal{A}$ for all other TSM names (i.e. the previous mapping, if it exists, is updated).

A *ueTSM definition context*, Ψ, is a finite function mapping each symbol $a \in \text{dom}(\Psi)$ to an *expanded ueTSM definition*, $a \hookrightarrow \texttt{uetsm}(\tau; e_{\text{parse}})$, where τ is the ueTSM's type

---

[1]In VerseML, the `ParseError` constructor of `ParseResult` required an error message and an error location, but we omit these in our formalization for simplicity

annotation, and e_{parse} is its parse function. We write $\Psi, a \hookrightarrow \texttt{uetsm}(\tau; e_{\text{parse}})$ when $a \notin \text{dom}(\Psi)$ for the extension of Ψ that maps a to $a \hookrightarrow \texttt{uetsm}(\tau; e_{\text{parse}})$. We write $\Delta \vdash \Psi$ ueTSMs when all the type annotations in Ψ are well-formed assuming Δ, and the parse functions in Ψ are closed and of type $\texttt{Body} \rightharpoonup \texttt{ParseResultExp}$.

Definition 3.17 (ueTSM Definition Context Formation). $\Delta \vdash \Psi$ ueTSMs *iff for each* $\hat{a} \hookrightarrow \textbf{\textit{uetsm}}(\tau; e_{\textit{parse}}) \in \Psi$, *we have* $\Delta \vdash \tau$ type *and* $\emptyset \emptyset \vdash e_{\textit{parse}} : \texttt{Body} \rightharpoonup \texttt{ParseResultExp}$.

We define $\hat{\Psi}, \hat{a} \rightsquigarrow a \hookrightarrow \texttt{uetsm}(\tau; e_{\text{parse}})$, when $\hat{\Psi} = \langle \mathcal{A}; \Psi \rangle$, as an abbreviation of

$$\langle \mathcal{A} \uplus \hat{a} \rightsquigarrow a; \Psi, a \hookrightarrow \texttt{uetsm}(\tau; e_{\text{parse}}) \rangle$$

The reason we distinguish TSM names, $\hat{a}$, from symbols, a, is in order to support TSM abbreviations, which we defer discussion of until Chapter 5.

3.2.7 ueTSM Application

The stylized unexpanded expression form for applying a ueTSM named $\hat{a}$ to a literal form with literal body b is:

$$\hat{a} \ /b/$$

This stylized form uses forward slashes to delimit the literal body, but stylized variants of any of the literal forms specified in Figure 3.1 could also be added to Figure 3.5. The corresponding operational form is $\texttt{uapuetsm}[b][\hat{a}]$.

The typed expansion rule governing ueTSM application is below:

$$
\frac{
\begin{array}{c}
b \downarrow e_{\text{body}} \qquad e_{\text{parse}}(e_{\text{body}}) \Downarrow \texttt{inj Success} \cdot e_{\text{cand}} \qquad e_{\text{cand}} \uparrow_{\text{CEExp}} \grave{e} \\
\emptyset \emptyset \vdash^{\hat{\Delta}; \hat{\Gamma}; \hat{\Psi}, \hat{a} \rightsquigarrow a \hookrightarrow \texttt{uetsm}(\tau; e_{\text{parse}}); b} \grave{e} \rightsquigarrow e : \tau
\end{array}
}{
\hat{\Delta} \ \hat{\Gamma} \vdash_{\hat{\Psi}, \hat{a} \rightsquigarrow a \hookrightarrow \texttt{uetsm}(\tau; e_{\text{parse}})} \hat{a} \ /b/ \rightsquigarrow e : \tau
} \tag{3.4m}
$$

The premises of Rule (3.4m) can be understood as follows, in order:

1. The first premise determines the encoding of the literal body, e_{body} (see above).

2. The second premise applies the parse function e_{parse}, which appears in the ueTSM context associated with $\hat{a}$, to e_{body}. If parsing succeeds, i.e. a value of the (stylized) form $\texttt{inj Success} \cdot e_{\text{cand}}$ results from evaluation, then e_{cand} will be a value of type $\texttt{CEExp}$ (assuming a well-formed ueTSM context, by application of Assumption 3.9). We call e_{cand} the *encoding of the candidate expansion*.

   If the parse function produces a value labeled $\texttt{ParseError}$, then typed expansion fails. No rule is necessary to handle this case.

3. The third premise decodes the encoding of the candidate expansion to produce the *candidate expansion*, $\grave{e}$ (see above).

4. The final premise of Rule (3.4m) *validates* the candidate expansion and simultaneously generates the *final expansion*, e. This is the topic of Sec. 3.2.9.

Sort	Operational Form	Stylized Form	Description
CETyp $\grave{\tau}$ ::=	t	t	variable
	$\texttt{ceparr}(\grave{\tau};\grave{\tau})$	$\grave{\tau} \rightharpoonup \grave{\tau}$	partial function
	$\texttt{ceall}(t.\grave{\tau})$	$\forall t.\grave{\tau}$	polymorphic
	$\texttt{cerec}(t.\grave{\tau})$	$\mu t.\grave{\tau}$	recursive
	$\texttt{ceprod}[L](\{i \hookrightarrow \grave{\tau}_i\}_{i \in L})$	$\langle\{i \hookrightarrow \grave{\tau}_i\}_{i \in L}\rangle$	labeled product
	$\texttt{cesum}[L](\{i \hookrightarrow \grave{\tau}_i\}_{i \in L})$	$[\{i \hookrightarrow \grave{\tau}_i\}_{i \in L}]$	labeled sum
	$\texttt{cesplicedt}[m;n]$	$\texttt{splicedt}\langle m,n \rangle$	spliced
CEExp $\grave{e}$ ::=	x	x	variable
	$\texttt{celam}\{\grave{\tau}\}(x.\grave{e})$	$\lambda x{:}\grave{\tau}.\grave{e}$	abstraction
	$\texttt{ceap}(\grave{e};\grave{e})$	$\grave{e}(\grave{e})$	application
	$\texttt{cetlam}(t.\grave{e})$	$\Lambda t.\grave{e}$	type abstraction
	$\texttt{cetap}\{\grave{\tau}\}(\grave{e})$	$\grave{e}[\grave{\tau}]$	type application
	$\texttt{cefold}\{t.\grave{\tau}\}(\grave{e})$	$\texttt{fold}(\grave{e})$	fold
	$\texttt{ceunfold}(\grave{e})$	$\texttt{unfold}(\grave{e})$	unfold
	$\texttt{cetpl}[L](\{i \hookrightarrow \grave{e}_i\}_{i \in L})$	$\langle\{i \hookrightarrow \grave{e}_i\}_{i \in L}\rangle$	labeled tuple
	$\texttt{cepr}[\ell](\grave{e})$	$\grave{e} \cdot \ell$	projection
	$\texttt{cein}[L][\ell]\{\{i \hookrightarrow \grave{\tau}_i\}_{i \in L}\}(\grave{e})$	$\texttt{inj}\ \ell \cdot \grave{e}$	injection
	$\texttt{cecase}[L]\{\tau\}(\grave{e};\{i \hookrightarrow x_i.\grave{e}_i\}_{i \in L})$	$\texttt{case}\ \grave{e}\ \{i \hookrightarrow x_i.\grave{e}_i\}_{i \in L}$	case analysis
	$\texttt{cesplicede}[m;n]$	$\texttt{splicede}\langle m,n \rangle$	spliced

Figure 3.6: Abstract syntax of candidate expansion types and expressions in miniVerse$_{\textbf{UE}}$. Metavariables m and n range over natural numbers. Candidate expansion types and expressions are identified up to α-equivalence.

3.2.8 Syntax of Candidate Expansions

Figure 3.6 defines the syntax of candidate expansion types (or *ce-types*), $\grave{\tau}$, and candidate expansion expressions (or *ce-expressions*), $\grave{e}$. Candidate expansion types and expressions are identified up to α-equivalence in the usual manner.

Each inner core form maps onto a candidate expansion form. We refer to these as the *shared forms*. In particular:

- Each type form maps onto a ce-type form according to the metafunction $\mathcal{C}(\tau)$, defined as follows:

$$
\begin{aligned}
\mathcal{C}(t) &= t \\
\mathcal{C}(\texttt{parr}(\tau_1;\tau_2)) &= \texttt{ceparr}(\mathcal{C}(\tau_1);\mathcal{C}(\tau_2)) \\
\mathcal{C}(\texttt{all}(t.\tau)) &= \texttt{ceall}(t.\mathcal{C}(\tau)) \\
\mathcal{C}(\texttt{rec}(t.\tau)) &= \texttt{cerec}(t.\mathcal{C}(\tau)) \\
\mathcal{C}(\texttt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L})) &= \texttt{ceprod}[L](\{i \hookrightarrow \mathcal{C}(\grave{\tau}_i)\}_{i \in L}) \\
\mathcal{C}(\texttt{sum}[L](\{i \hookrightarrow \tau_i\}_{i \in L})) &= \texttt{cesum}[L](\{i \hookrightarrow \mathcal{C}(\grave{\tau}_i)\}_{i \in L})
\end{aligned}
$$

- Each expanded expression form maps onto a ce-expression form according to the

metafunction $\mathcal{C}(e)$, defined as follows:

$$\mathcal{C}(x) = x$$
$$\mathcal{C}(\texttt{lam}\{\tau\}(x.e)) = \texttt{celam}\{\mathcal{C}(\tau)\}(x.\mathcal{C}(e))$$
$$\mathcal{C}(\texttt{ap}(e_1; e_2)) = \texttt{ceap}(\mathcal{C}(e_1); \mathcal{C}(e_2))$$
$$\mathcal{C}(\texttt{tlam}(t.e)) = \texttt{cetlam}(t.\mathcal{C}(e))$$
$$\mathcal{C}(\texttt{tap}\{\tau\}(e)) = \texttt{cetap}\{\mathcal{C}(\tau)\}(\mathcal{C}(e))$$
$$\mathcal{C}(\texttt{fold}\{t.\tau\}(e)) = \texttt{cefold}\{t.\mathcal{C}(\tau)\}(\mathcal{C}(e))$$
$$\mathcal{C}(\texttt{unfold}(e)) = \texttt{ceunfold}(\mathcal{C}(e))$$
$$\mathcal{C}(\texttt{tpl}[L](\{i \hookrightarrow e_i\}_{i \in L})) = \texttt{cetpl}[L](\{i \hookrightarrow \mathcal{C}(e_i)\}_{i \in L})$$
$$\mathcal{C}(\texttt{in}[L][\ell]\{\{i \hookrightarrow \tau_i\}_{i \in L}\}(e)) = \texttt{cein}[L][\ell]\{\{i \hookrightarrow \mathcal{C}(\tau_i)\}_{i \in L}\}(\mathcal{C}(e))$$
$$\mathcal{C}(\texttt{case}[L]\{\tau\}(e; \{i \hookrightarrow x_i.e_i\}_{i \in L})) = \texttt{cecase}[L]\{\mathcal{C}(\tau)\}(\mathcal{C}(e); \{i \hookrightarrow x_i.\mathcal{C}(e_i)\}_{i \in L})$$

There are two other candidate expansion forms, highlighted in gray in Figure 3.6: a ce-type form for *references to spliced unexpanded types*, $\texttt{cesplicedt}[m; n]$, and a ce-expression form for *references to spliced unexpanded expressions*, $\texttt{cesplicede}[m; n]$.

3.2.9 Candidate Expansion Validation

The *candidate expansion validation judgements* validate ce-types and ce-expressions and simultaneously generate their final expansions.

Judgement Form	Description
$\Delta \vdash^{\mathbb{T}} \check{\tau} \rightsquigarrow \tau$ type	Candidate expansion type $\check{\tau}$ is well-formed and has expansion τ assuming Δ and type splicing scene $\mathbb{T}$.
$\Delta\,\Gamma \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e : \tau$	Candidate expansion expression $\grave{e}$ has expansion e and type τ assuming Δ and Γ and expression splicing scene $\mathbb{E}$.

Expression splicing scenes, $\mathbb{E}$, are of the form $\hat{\Delta}$; $\hat{\Gamma}$; $\hat{\Psi}$; b, and *type splicing scenes*, $\mathbb{T}$, are of the form Δ; b. We write $\texttt{ts}(\mathbb{E})$ for the type splicing scene constructed by dropping the unexpanded typing context and ueTSM context from $\mathbb{E}$:

$$\texttt{ts}(\hat{\Delta}; \hat{\Gamma}; \hat{\Psi}; b) = \hat{\Delta}; b$$

The purpose of splicing scenes is to "remember", during the candidate expansion validation process, the unexpanded type formation context, $\hat{\Delta}$, unexpanded typing context, $\hat{\Gamma}$, ueTSM context, $\hat{\Psi}$, and the literal body, b, from the ueTSM application site (cf. Rule (3.4m)), because these are necessary to validate references to spliced unexpanded types and expressions that appear within a candidate expansion.

Candidate Expansion Type Validation

The *candidate expansion type validation judgement*, $\Delta \vdash^{\mathbb{T}} \check{\tau} \rightsquigarrow \tau$ type, is inductively defined by Rules (3.5) as follows.

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Shared Forms   Rules (3.5a) through (3.5f), which validate ce-types of shared form, are defined below.

$$\frac{}{\Delta, t \text{ type} \vdash^{\mathbb{T}} t \rightsquigarrow t \text{ type}} \tag{3.5a}$$

$$\frac{\Delta \vdash^{\mathbb{T}} \tilde{\tau}_1 \rightsquigarrow \tau_1 \text{ type} \qquad \Delta \vdash^{\mathbb{T}} \tilde{\tau}_2 \rightsquigarrow \tau_2 \text{ type}}{\Delta \vdash^{\mathbb{T}} \texttt{ceparr}(\tilde{\tau}_1; \tilde{\tau}_2) \rightsquigarrow \texttt{parr}(\tau_1; \tau_2) \text{ type}} \tag{3.5b}$$

$$\frac{\Delta, t \text{ type} \vdash^{\mathbb{T}} \tilde{\tau} \rightsquigarrow \tau \text{ type}}{\Delta \vdash^{\mathbb{T}} \texttt{ceall}(t.\tilde{\tau}) \rightsquigarrow \texttt{all}(t.\tau) \text{ type}} \tag{3.5c}$$

$$\frac{\Delta, t \text{ type} \vdash^{\mathbb{T}} \tilde{\tau} \rightsquigarrow \tau \text{ type}}{\Delta \vdash^{\mathbb{T}} \texttt{cerec}(t.\tilde{\tau}) \rightsquigarrow \texttt{rec}(t.\tau) \text{ type}} \tag{3.5d}$$

$$\frac{\{\Delta \vdash^{\mathbb{T}} \tilde{\tau}_i \rightsquigarrow \tau_i \text{ type}\}_{i \in L}}{\Delta \vdash^{\mathbb{T}} \texttt{ceprod}[L](\{i \hookrightarrow \tilde{\tau}_i\}_{i \in L}) \rightsquigarrow \texttt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L}) \text{ type}} \tag{3.5e}$$

$$\frac{\{\Delta \vdash^{\mathbb{T}} \tilde{\tau}_i \rightsquigarrow \tau_i \text{ type}\}_{i \in L}}{\Delta \vdash^{\mathbb{T}} \texttt{cesum}[L](\{i \hookrightarrow \tilde{\tau}_i\}_{i \in L}) \rightsquigarrow \texttt{sum}[L](\{i \hookrightarrow \tau_i\}_{i \in L}) \text{ type}} \tag{3.5f}$$

Observe that, in each of these rules, the ce-type form and the type form in the conclusion correspond, and the premises correspond to those of the corresponding type formation rule, i.e. Rules (3.1). The type splicing scene, $\mathbb{T}$, passes opaquely through these rules. The following lemma establishes that each type can be expressed as a well-formed ce-type, under the same type formation context and any type splicing scene.

Lemma 3.18 (Candidate Expansion Type Expressibility)**.** *If* $\Delta \vdash \tau$ type *then* $\Delta \vdash^{\mathbb{T}} \mathcal{C}(\tau) \rightsquigarrow \tau$ type*.*

Proof. By rule induction over Rules (3.1). In each case, we apply the IH on or over each premise, then apply the corresponding ce-type validation rule in Rules (3.5). □

Notice that in Rule (3.5a), only type variables tracked by the candidate expansion type formation context, Δ, are validated. Type variables in the application site unexpanded type formation context, which appears within the type splicing scene, $\mathbb{T}$, are not validated. Indeed, $\mathbb{T}$ is not inspected by any of the rules above. This achieves *context-independent expansion* as described in Sec. 3.1.3 for type variables – ueTSMs cannot impose "hidden constraints" on the application site unexpanded type formation context, because the type variables bound at the application site are simply not directly available to ce-types.

References to Spliced Types   The only ce-type form that does not correspond to a type form is `cesplicedt[m; n]`, which is a *reference to a spliced unexpanded type*, i.e. it indicates that an unexpanded type should be parsed out from the literal body, b, which appears in the type splicing scene, beginning at position m and ending at position n.

Rule (3.5g) governs this form:

$$\frac{\mathsf{parseUTyp}(\mathsf{subseq}(b;m;n)) = \hat{\tau} \qquad \langle \mathcal{D}; \Delta_{\mathrm{app}} \rangle \vdash \hat{\tau} \rightsquigarrow \tau \text{ type} \qquad \Delta \cap \Delta_{\mathrm{app}} = \varnothing}{\Delta \vdash^{\langle \mathcal{D}; \Delta_{\mathrm{app}} \rangle; b} \texttt{cesplicedt}[m;n] \rightsquigarrow \tau \text{ type}} \tag{3.5g}$$

The first premise of this rule extracts the indicated subsequence of b using the partial metafunction $\mathsf{subseq}(b;m;n)$ and parses it using the partial metafunction $\mathsf{parseUTyp}(b)$, described in Sec. 3.2.4, to produce the spliced unexpanded type itself, $\hat{\tau}$.

The second premise of Rule (3.5g) performs type expansion of $\hat{\tau}$ under the application site unexpanded type formation context, $\langle \mathcal{D}; \Delta_{\mathrm{app}} \rangle$, which appears in the type splicing scene. The hypotheses in the candidate expansion type formation context, Δ, are not made available to τ.

The third premise of Rule (3.5g) imposes the constraint that the candidate expansion's type formation context, Δ, be disjoint from the application site type formation context, Δ_{app}. This premise can always be discharged by α-varying the candidate expansion that the reference to the spliced type appears within.

This achieves *expansion-independent splicing* as described in Sec. 3.1.3 for type variables – the TSM provider can choose type variable names freely within a candidate expansion, because the language prevents them from shadowing type variables at the application site (by α-varying the candidate expansion as needed).

Rules (3.5) validate the following lemma, which establishes that the final expansion of a valid ce-type is a well-formed type under the combined type formation context.

Lemma 3.19 (Candidate Expansion Type Validation). *If* $\Delta \vdash^{\langle \mathcal{D}; \Delta_{app} \rangle; b} \hat{\tau} \rightsquigarrow \tau \text{ type}$ *then* $\Delta \cup \Delta_{app} \vdash \tau \text{ type}$.

Proof. By rule induction over Rules (3.5).

Case (3.5a). We have

(1)	$\Delta = \Delta', t \text{ type}$	by assumption
(2)	$\hat{\tau} = t$	by assumption
(3)	$\tau = t$	by assumption
(4)	$\Delta', t \text{ type} \vdash t \text{ type}$	by Rule (3.1a)
(5)	$\Delta', t \text{ type} \cup \Delta_{\mathrm{app}} \vdash t \text{ type}$	by Lemma 3.2 over Δ_{app} to (4)

Case (3.5b). We have

(1)	$\hat{\tau} = \texttt{ceparr}(\hat{\tau}_1; \hat{\tau}_2)$	by assumption
(2)	$\tau = \texttt{parr}(\tau_1; \tau_2)$	by assumption
(3)	$\Delta \vdash^{\langle \mathcal{D}; \Delta_{\mathrm{app}} \rangle; b} \hat{\tau}_1 \rightsquigarrow \tau_1 \text{ type}$	by assumption
(4)	$\Delta \vdash^{\langle \mathcal{D}; \Delta_{\mathrm{app}} \rangle; b} \hat{\tau}_2 \rightsquigarrow \tau_2 \text{ type}$	by assumption
(5)	$\Delta \cup \Delta_{\mathrm{app}} \vdash \tau_1 \text{ type}$	by IH on (3)
(6)	$\Delta \cup \Delta_{\mathrm{app}} \vdash \tau_2 \text{ type}$	by IH on (4)

(7) $\Delta \cup \Delta_{\mathrm{app}} \vdash \mathtt{parr}(\tau_1; \tau_2)$ type

by Rule (3.1b) on (5) and (6)

Case (3.5c). We have

(1) $\hat{\tau} = \mathtt{ceall}(t.\hat{\tau}')$ by assumption
(2) $\tau = \mathtt{all}(t.\tau')$ by assumption
(3) $\Delta, t \text{ type} \vdash^{\langle \mathcal{D}; \Delta_{\mathrm{app}} \rangle; b} \hat{\tau}' \rightsquigarrow \tau'$ type   by assumption
(4) $\Delta, t \text{ type} \cup \Delta_{\mathrm{app}} \vdash \tau'$ type   by IH on (3)
(5) $\Delta \cup \Delta_{\mathrm{app}}, t \text{ type} \vdash \tau'$ type   by exchange over Δ_{app} on (4)
(6) $\Delta \cup \Delta_{\mathrm{app}} \vdash \mathtt{all}(t.\tau')$ type   by Rule (3.1c) on (5)

Case (3.5d) through (3.5f). These cases follow analagously, i.e. we apply the IH to or over all ce-type validation premises, apply exchange as needed, and then apply the corresponding type formation rule.

Case (3.5g). We have

(1) $\hat{\tau} = \mathtt{cesplicedt}[m; n]$ by assumption
(2) $\mathsf{parseUTyp}(\mathsf{subseq}(b; m; n)) = \hat{\tau}$ by assumption
(3) $\langle \mathcal{D}; \Delta_{\mathrm{app}} \rangle \vdash \hat{\tau} \rightsquigarrow \tau$ type   by assumption
(4) $\Delta_{\mathrm{app}} \vdash \tau$ type   by Lemma 3.12 on (3)
(5) $\Delta \cup \Delta_{\mathrm{app}} \vdash \tau$ type   by Lemma 3.2 over Δ on (4) and exchange over Δ

$\square$

Candidate Expansion Expression Validation

The *candidate expansion expression validation judgement*, $\Delta\ \Gamma \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e : \tau$, is defined mutually inductively with the typed expansion judgement by Rules (3.6) as follows.

Shared Forms   Rules (3.6a) through (3.6k) validate ce-expressions of shared form. The first three of these rules are defined below:

$$\frac{}{\Delta\ \Gamma, x : \tau \vdash^{\mathbb{E}} x \rightsquigarrow x : \tau} \tag{3.6a}$$

$$\frac{\Delta \vdash^{\mathsf{ts}(\mathbb{E})} \hat{\tau} \rightsquigarrow \tau \text{ type} \qquad \Delta\ \Gamma, x : \tau \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e : \tau'}{\Delta\ \Gamma \vdash^{\mathbb{E}} \mathtt{celam}\{\hat{\tau}\}(x.\grave{e}) \rightsquigarrow \mathtt{lam}\{\tau\}(x.e) : \mathtt{parr}(\tau; \tau')} \tag{3.6b}$$

$$\frac{\Delta\ \Gamma \vdash^{\mathbb{E}} \grave{e}_1 \rightsquigarrow e_1 : \mathtt{parr}(\tau; \tau') \qquad \Delta\ \Gamma \vdash^{\mathbb{E}} \grave{e}_2 \rightsquigarrow e_2 : \tau}{\Delta\ \Gamma \vdash^{\mathbb{E}} \mathtt{ceap}(\grave{e}_1; \grave{e}_2) \rightsquigarrow \mathtt{ap}(e_1; e_2) : \tau'} \tag{3.6c}$$

Observe that, in each of these rules, the ce-expression form and the expanded expression form in the conclusion correspond, and the premises correspond to those of the corresponding typing rule, i.e. Rules (3.2a) through (3.2c), respectively. The expression splicing scene, $\mathbb{E}$, passes opaquely through these rules.

We can express this scheme more precisely with the following rule transformation. For each rule in Rules (3.2),

$$\frac{J_1 \quad \cdots \quad J_k}{J}$$

the corresponding candidate expansion expression validation rule is

$$\frac{\mathcal{C}(J_1) \quad \cdots \quad \mathcal{C}(J_k)}{\mathcal{C}(J)}$$

where

$$\mathcal{C}(\Delta \vdash \tau \text{ type}) = \Delta \vdash^{\text{ts}(\mathbb{E})} \mathcal{C}(\tau) \rightsquigarrow \tau \text{ type}$$
$$\mathcal{C}(\Delta \, \Gamma \vdash e : \tau) = \Delta \, \Gamma \vdash^{\mathbb{E}} \mathcal{C}(e) \rightsquigarrow e : \tau$$
$$\mathcal{C}(\{J_i\}_{i \in L}) = \{\mathcal{C}(J_i)\}_{i \in L}$$

and where:
- $\mathcal{C}(\tau)$ is defined as follows:
  - When τ is of definite form, $\mathcal{C}(\tau)$ is defined as in Sec. 3.2.8.
  - When τ is of indefinite form, $\mathcal{C}(\tau)$ is a uniquely corresponding metavariable of sort CETyp also of indefinite form. For example, $\mathcal{C}(\tau_1) = \hat{\tau}_1$ and $\mathcal{C}(\tau_2) = \hat{\tau}_2$.
- $\mathcal{C}(e)$ is defined as follows
  - When e is of definite form, $\mathcal{C}(e)$ is defined as in Sec. 3.2.8.
  - When e is of indefinite form, $\mathcal{C}(e)$ is a uniquely corresponding metavariable of sort CEExp also of indefinite form. For example, $\mathcal{C}(e_1) = \hat{e}_1$ and $\mathcal{C}(e_2) = \hat{e}_2$.

It is instructive to use this rule transformation to generate Rules (3.6a) through (3.6c) above. We omit the remaining rules for shared forms, i.e. Rules (3.6d) through (3.6k).

The following lemma establishes that each well-typed expanded expression, e, can be expressed as a valid ce-expression, $\mathcal{C}(e)$, that is assigned the same type under any expression splicing scene.

Theorem 3.20 (Candidate Expansion Expression Expressibility)**.** *If $\Delta \, \Gamma \vdash e : \tau$ then $\Delta \, \Gamma \vdash^{\mathbb{E}} \mathcal{C}(e) \rightsquigarrow e : \tau$.*

Proof. By rule induction over Rules (3.2). The rule transformation above guarantees that this lemma holds by construction. In particular, in each case, we apply Lemma 3.18 to or over each type formation premise, the IH to or over each typing premise, then apply the corresponding ce-expression validation rule in Rules (3.6a) through (3.6k). □

Notice that in Rule (3.6a), only variables tracked by the candidate expansion typing context, Γ, are validated. Variables in the application site unexpanded typing context, which appears within the expression splicing scene $\mathbb{E}$, are not validated. Indeed, $\mathbb{E}$ is not inspected by any of the rules above. This achieves *context-independent expansion* as described in Sec. 3.1.3 – ueTSMs cannot impose "hidden constraints" on the application site unexpanded typing context, because the variable bindings at the application site are not directly available to candidate expansions.

References to Spliced Unexpanded Expressions The only ce-expression form that does not correspond to an expanded expression form is $\mathtt{cesplicede}[m;n]$, which is a *reference to a spliced unexpanded expression*, i.e. it indicates that an unexpanded expression should be parsed out from the literal body beginning at position m and ending at position n. Rule (3.6l) governs this form:

$$\frac{\mathsf{parseUExp}(\mathsf{subseq}(b;m;n)) = \hat{e} \qquad \langle \mathcal{D};\Delta_{\mathsf{app}}\rangle \, \langle \mathcal{G};\Gamma_{\mathsf{app}}\rangle \vdash_{\hat{\Psi}} \hat{e} \rightsquigarrow e : \tau \\ \Delta \cap \Delta_{\mathsf{app}} = \varnothing \qquad \mathsf{dom}(\Gamma) \cap \mathsf{dom}(\Gamma_{\mathsf{app}}) = \varnothing}{\Delta \, \Gamma \vdash^{\langle \mathcal{D};\Delta_{\mathsf{app}}\rangle ; \langle \mathcal{G};\Gamma_{\mathsf{app}}\rangle ; \hat{\Psi} ; b} \mathtt{cesplicede}[m;n] \rightsquigarrow e : \tau} \tag{3.6l}$$

The first premise of this rule extracts the indicated subsequence of b using the partial metafunction $\mathsf{subseq}(b;m;n)$ and parses it using the partial metafunction $\mathsf{parseUExp}(b)$, described in Sec. 3.2.4, to produce the referenced spliced unexpanded expression, $\hat{e}$.

The second premise of Rule (3.6l) types and expands the spliced unexpanded expression $\hat{e}$ assuming the application site contexts that appear in the expression splicing scene. The hypotheses in the candidate expansion type formation context, Δ, and typing context, Γ, are not made available to $\hat{e}$.

The third premise of Rule (3.6l) imposes the constraint that the candidate expansion's type formation context, Δ, be disjoint from the application site type formation context, Δ_{app}. Similarly, the fourth premise requires that the candidate expansion's typing context, Γ, be disjoint from the application site typing context, Γ_{app}. These two premises can always be discharged by α-varying the ce-expression that the reference to the spliced unexpanded expression appears within.

This achieves *expansion-independent splicing* as described in Sec. 3.1.3 – the TSM provider can choose variable names freely within a candidate expansion, because the language prevents them from shadowing those at the application site (by α-varying the candidate expansion as needed).

3.2.10 Metatheory

For the judgements we have defined to form a sensible language, we must have that typed expansion and candidate expansion expression validation be consistent with typing. Formally, this can be expressed as follows.

Theorem 3.21 (Typed Expansion). *Both of the following hold:*

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 $\Delta \cap \Delta_{app} = \varnothing$ and $dom(\Gamma) \cap dom(\Gamma_{app}) = \varnothing$ then $\Delta \cup \Delta_{app}\,\Gamma \cup \Gamma_{app} \vdash e : \tau$.

Proof. By mutual rule induction over Rules (3.4) and Rules (3.6).

The proof of part 1 proceeds by inducting over the typed expansion assumption. In the following cases, let $\hat{\Delta} = \langle\mathcal{D};\Delta\rangle$ and $\hat{\Gamma} = \langle\mathcal{G};\Gamma\rangle$ and $\hat{\Psi} = \langle\mathcal{A};\Psi\rangle$.

Case (3.4a). We have:

(1)	$\hat{e} = \hat{x}$	by assumption
(2)	$e = x$	by assumption
(3)	$\Gamma = \Gamma', x : \tau$	by assumption
(4)	$\Delta\,\Gamma', x : \tau \vdash x : \tau$	by Rule (3.2a)

Case (3.4b). We have:

(1)	$\hat{e} = \mathtt{ulam}\{\hat{\tau}_1\}(\hat{x}.\hat{e}')$	by assumption
(2)	$e = \mathtt{lam}\{\tau_1\}(x.e')$	by assumption
(3)	$\tau = \mathtt{parr}(\tau_1;\tau_2)$	by assumption
(4)	$\hat{\Delta} \vdash \hat{\tau}_1 \rightsquigarrow \tau_1\ \mathsf{type}$	by assumption
(5)	$\hat{\Delta}\,\hat{\Gamma}, \hat{x} \rightsquigarrow x : \tau_1 \vdash_{\hat{\Psi}} \hat{e}' \rightsquigarrow e' : \tau_2$	by assumption
(6)	$\Delta \vdash \tau_1\ \mathsf{type}$	by Lemma 3.12 on (4)
(7)	$\Delta\,\Gamma, x : \tau_1 \vdash e' : \tau_2$	by IH, part 1 on (5)
(8)	$\Delta\,\Gamma \vdash \mathtt{lam}\{\tau_1\}(x.e') : \mathtt{parr}(\tau_1;\tau_2)$	by Rule (3.2b) on (6) and (7)

Case (3.4c). We have:

(1)	$\hat{e} = \mathtt{uap}(\hat{e}_1;\hat{e}_2)$	by assumption
(2)	$e = \mathtt{ap}(e_1;e_2)$	by assumption
(3)	$\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi}} \hat{e}_1 \rightsquigarrow e_1 : \mathtt{parr}(\tau_1;\tau)$	by assumption
(4)	$\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi}} \hat{e}_2 \rightsquigarrow e_2 : \tau_1$	by assumption
(5)	$\Delta\,\Gamma \vdash e_1 : \mathtt{parr}(\tau_1;\tau)$	by IH, part 1 on (3)
(6)	$\Delta\,\Gamma \vdash e_2 : \tau_1$	by IH on (4)
(7)	$\Delta\,\Gamma \vdash \mathtt{ap}(e_1;e_2) : \tau$	by Rule (3.2c) on (5) and (6)

Case (3.4d) through (3.4k). These cases follow analagously, i.e. we apply Lemma 3.12 to or over the type expansion premises and the IH, part 1, to or over the typed expression expansion premises and then apply the corresponding typing rule in Rules (3.2d) through (3.2k).

Case (3.4l). We have

(1)	$\hat{e} = \mathtt{usyntaxue}\{e_{\mathrm{parse}}\}\{\hat{\tau}'\}(\hat{a}.\hat{e}')$	by assumption

(2) $\hat{\Delta} \vdash \hat{\tau}' \rightsquigarrow \tau'$ type   by assumption

(3) $\hat{\Delta}\,\hat{\Gamma} \vdash_{\Psi,\hat{a}\rightsquigarrow a\hookrightarrow \texttt{uetsm}(\tau';e_{\text{parse}})} \hat{e}' \rightsquigarrow e : \tau$ by assumption

(4) $\Delta\,\Gamma \vdash e : \tau$ by IH, part 1 on (3)

Case (3.4m). We have

    (1) $\hat{e} = \texttt{uapuetsm}[b][\hat{a}]$ by assumption

    (2) $\mathcal{A} = \mathcal{A}', \hat{a} \rightsquigarrow a$ by assumption

    (3) $\Psi = \Psi', a \hookrightarrow \texttt{uetsm}(\tau; e_{\text{parse}})$ by assumption

    (4) $b \downarrow e_{\text{body}}$ by assumption

    (5) $e_{\text{parse}}(e_{\text{body}}) \Downarrow \texttt{inj Success} \cdot e_{\text{cand}}$ by assumption

    (6) $e_{\text{cand}} \uparrow_{\text{CEExp}} \grave{e}$ by assumption

    (7) $\varnothing\,\varnothing \vdash^{\hat{\Delta};\hat{\Gamma};\hat{\Psi};b} \grave{e} \rightsquigarrow e : \tau$ by assumption

    (8) $\varnothing \cap \Delta = \varnothing$ by finite set intersection identity

    (9) $\varnothing \cap \text{dom}(\Gamma) = \varnothing$ by finite set intersection identity

   (10) $\varnothing \cup \Delta\,\varnothing \cup \Gamma \vdash e : \tau$ by IH, part 2 on (7), (8), and (9)

   (11) $\Delta\,\Gamma \vdash e : \tau$ by definition of finite set and finite function union over (10)

The second part of the theorem proceeds by induction over the candidate expansion expression validation assumption as follows. In the following cases, let $\hat{\Delta}_{\text{app}} = \langle \mathcal{D}; \Delta_{\text{app}} \rangle$ and $\hat{\Gamma}_{\text{app}} = \langle \mathcal{G}; \Gamma_{\text{app}} \rangle$ and $\hat{\Psi} = \langle \mathcal{A}; \Psi \rangle$.

Case (3.6a). We have

    (1) $\grave{e} = x$ by assumption

    (2) $e = x$ by assumption

    (3) $\Gamma = \Gamma', x : \tau$ by assumption

    (4) $\Delta \cup \Delta_{\text{app}}\,\Gamma', x : \tau \vdash x : \tau$ by Rule (3.2a)

    (5) $\Delta \cup \Delta_{\text{app}}\,\Gamma', x : \tau \cup \Gamma_{\text{app}} \vdash x : \tau$ by Lemma 3.2 over Γ_{app} to (4)

Case (3.6b). We have

    (1) $\grave{e} = \texttt{celam}\{\grave{\tau}_1\}(x.\grave{e}')$ by assumption

    (2) $e = \texttt{lam}\{\tau_1\}(x.e')$ by assumption

    (3) $\tau = \texttt{parr}(\tau_1; \tau_2)$ by assumption

    (4) $\Delta \vdash^{\hat{\Delta}_{\text{app}};b} \grave{\tau}_1 \rightsquigarrow \tau_1$ type   by assumption

    (5) $\Delta\,\Gamma, x : \tau_1 \vdash^{\hat{\Delta}_{\text{app}};\hat{\Gamma}_{\text{app}};\hat{\Psi};b} \grave{e}' \rightsquigarrow e' : \tau_2$ by assumption

    (6) $\Delta \cap \Delta_{\text{app}} = \varnothing$ by assumption

(7)	$\mathrm{dom}(\Gamma) \cap \mathrm{dom}(\Gamma_{\mathrm{app}}) = \varnothing$	by assumption
(8)	$x \notin \mathrm{dom}(\Gamma_{\mathrm{app}})$	by identification convention
(9)	$\mathrm{dom}(\Gamma, x : \tau_1) \cap \mathrm{dom}(\Gamma_{\mathrm{app}}) = \varnothing$	by (7) and (8)
(10)	$\Delta \cup \Delta_{\mathrm{app}} \vdash \tau_1 \; \mathrm{type}$	by Lemma 3.19 on (4)
(11)	$\Delta \cup \Delta_{\mathrm{app}} \; \Gamma, x : \tau_1 \cup \Gamma_{\mathrm{app}} \vdash e' : \tau_2$	by IH, part 2 on (5), (6) and (9)
(12)	$\Delta \cup \Delta_{\mathrm{app}} \; \Gamma \cup \Gamma_{\mathrm{app}}, x : \tau_1 \vdash e' : \tau_2$	by exchange over Γ_{app} on (11)
(13)	$\Delta \cup \Delta_{\mathrm{app}} \; \Gamma \cup \Gamma_{\mathrm{app}} \vdash \mathtt{lam}\{\tau_1\}(x.e') : \mathtt{parr}(\tau_1; \tau_2)$	by Rule (3.2b) on (10) and (12)

Case (3.6c). We have

(1)	$\grave{e} = \mathtt{ceap}(\grave{e}_1; \grave{e}_2)$	by assumption
(2)	$e = \mathtt{ap}(e_1; e_2)$	by assumption
(3)	$\Delta \; \Gamma \vdash^{\hat{\Delta}_{\mathrm{app}}; \hat{\Gamma}_{\mathrm{app}}; \hat{\Psi}; b} \grave{e}_1 \rightsquigarrow e_1 : \mathtt{parr}(\tau_2; \tau)$	by assumption
(4)	$\Delta \; \Gamma \vdash^{\hat{\Delta}_{\mathrm{app}}; \hat{\Gamma}_{\mathrm{app}}; \hat{\Psi}; b} \grave{e}_2 \rightsquigarrow e_2 : \tau_2$	by assumption
(5)	$\Delta \cap \Delta_{\mathrm{app}} = \varnothing$	by assumption
(6)	$\mathrm{dom}(\Gamma) \cap \mathrm{dom}(\Gamma_{\mathrm{app}}) = \varnothing$	by assumption
(7)	$\Delta \cup \Delta_{\mathrm{app}} \; \Gamma \cup \Gamma_{\mathrm{app}} \vdash e_1 : \mathtt{parr}(\tau_2; \tau)$	by IH, part 2 on (3), (5) and (6)
(8)	$\Delta \cup \Delta_{\mathrm{app}} \; \Gamma \cup \Gamma_{\mathrm{app}} \vdash e_2 : \tau_2$	by IH, part 2 on (4), (5) and (6)
(9)	$\Delta \cup \Delta_{\mathrm{app}} \; \Gamma \cup \Gamma_{\mathrm{app}} \vdash \mathtt{ap}(e_1; e_2) : \tau$	by Rule (3.2c) on (7) and (8)

Case (3.6d). We have

(1)	$\grave{e} = \mathtt{cetlam}(t.\grave{e}')$	by assumption
(2)	$e = \mathtt{tlam}(t.e')$	by assumption
(3)	$\tau = \mathtt{all}(t.\tau')$	by assumption
(4)	$\Delta, t \; \mathrm{type} \; \Gamma \vdash^{\hat{\Delta}_{\mathrm{app}}; \hat{\Gamma}_{\mathrm{app}}; \hat{\Psi}; b} \grave{e}' \rightsquigarrow e' : \tau'$	by assumption
(5)	$\Delta \cap \Delta_{\mathrm{app}} = \varnothing$	by assumption
(6)	$\mathrm{dom}(\Gamma) \cap \mathrm{dom}(\Gamma_{\mathrm{app}}) = \varnothing$	by assumption
(7)	$t \; \mathrm{type} \notin \Delta_{\mathrm{app}}$	by identification convention
(8)	$\Delta, t \; \mathrm{type} \cap \Delta_{\mathrm{app}} = \varnothing$	by (5) and (7)
(9)	$\Delta, t \; \mathrm{type} \cup \Delta_{\mathrm{app}} \; \Gamma \cup \Gamma_{\mathrm{app}} \vdash e' : \tau'$	by IH, part 2 on (4), (8) and (6)
(10)	$\Delta \cup \Delta_{\mathrm{app}}, t \; \mathrm{type} \; \Gamma \cup \Gamma_{\mathrm{app}} \vdash e' : \tau'$	by exchange over Δ_{app} on (9)
(11)	$\Delta \cup \Delta_{\mathrm{app}} \; \Gamma \cup \Gamma_{\mathrm{app}} \vdash \mathtt{tlam}(t.e') : \mathtt{all}(t.\tau')$	by Rule (3.2d) on (10)

Case (3.6e) **through** (3.6k)**.** These cases follow analagously, i.e. we apply the IH, part 2 to all ce-expression validation judgements, Lemma 3.19 to all ce-type validation judgements, the identification convention to ensure that extended contexts remain disjoint, weakening and exchange as needed, and the corresponding typing rule in Rules (3.2e) through (3.2k).

Case (3.6l)**.** We have

(1)	$\grave{e} = \mathtt{cesplicede}[m;n]$	by assumption
(2)	$\mathsf{parseUExp}(\mathsf{subseq}(b;m;n)) = \hat{e}$	by assumption
(3)	$\hat{\Delta}_{\mathrm{app}} \, \hat{\Gamma}_{\mathrm{app}} \vdash_{\hat{\Psi}} \hat{e} \rightsquigarrow e : \tau$	by assumption
(4)	$\Delta \cap \Delta_{\mathrm{app}} = \varnothing$	by assumption
(5)	$\mathrm{dom}(\Gamma) \cap \mathrm{dom}(\Gamma_{\mathrm{app}}) = \varnothing$	by assumption
(6)	$\Delta_{\mathrm{app}} \, \Gamma_{\mathrm{app}} \vdash e : \tau$	by IH, part 1 on (3)
(7)	$\Delta \cup \Delta_{\mathrm{app}} \, \Gamma \cup \Gamma_{\mathrm{app}} \vdash e : \tau$	by Lemma 3.2 over Δ and Γ and exchange on (6)

The mutual induction can be shown to be well-founded by showing that the following numeric metric on the judgements that we induct over is decreasing:

$$\|\hat{\Delta} \, \hat{\Gamma} \vdash_{\hat{\Psi}} \hat{e} \rightsquigarrow e : \tau\| = \|\hat{e}\|$$

$$\|\Delta \, \Gamma \vdash^{\hat{\Delta}_{\mathrm{app}}; \hat{\Gamma}_{\mathrm{app}}; \hat{\Psi}; b} \grave{e} \rightsquigarrow e : \tau\| = \|b\|$$

where $\|b\|$ is the length of b and $\|\hat{e}\|$ is the sum of the lengths of the literal bodies in $\hat{e}$,

$$\|\hat{x}\| = 0$$
$$\|\mathtt{ulam}\{\hat{\tau}\}(\hat{x}.\hat{e})\| = \|\hat{e}\|$$
$$\|\mathtt{uap}(\hat{e}_1; \hat{e}_2)\| = \|\hat{e}_1\| + \|\hat{e}_2\|$$
$$\|\mathtt{utlam}(\hat{t}.\hat{e})\| = \|\hat{e}\|$$
$$\|\mathtt{utap}\{\hat{\tau}\}(\hat{e})\| = \|\hat{e}\|$$
$$\|\mathtt{ufold}\{\hat{t}.\hat{\tau}\}(\hat{e})\| = \|\hat{e}\|$$
$$\|\mathtt{uunfold}(\hat{e})\| = \|\hat{e}\|$$
$$\|\mathtt{utpl}[L](\{i \hookrightarrow \hat{e}_i\}_{i \in L})\| = \sum_{i \in L} \|\hat{e}_i\|$$
$$\|\mathtt{upr}[\ell](\hat{e})\| = \|\hat{e}\|$$
$$\|\mathtt{uin}[L][\ell]\{\{i \hookrightarrow \hat{\tau}_i\}_{i \in L}\}(\hat{e})\| = \|\hat{e}\|$$
$$\|\mathtt{ucase}[L]\{\hat{\tau}\}(\hat{e}; \{i \hookrightarrow \hat{x}_i.\hat{e}_i\}_{i \in L})\| = \|\hat{e}\| + \sum_{i \in L} \|\hat{e}_i\|$$
$$\|\mathtt{usyntaxue}\{e_{\mathrm{parse}}\}\{\hat{\tau}\}(\hat{a}.\hat{e})\| = \|\hat{e}\|$$
$$\|\mathtt{uapuetsm}[b][\hat{a}]\| = \|b\|$$

The only case in the proof of part 1 that invokes part 2 is Case (3.4m). There, we have that the metric remains stable:

$$\|\hat{\Delta}\ \hat{\Gamma} \vdash_{\hat{\Psi}, \hat{a} \rightsquigarrow a \hookrightarrow \mathtt{uetsm}(\tau; e_{\mathrm{parse}})} \mathtt{uapuetsm}[b][\hat{a}] \rightsquigarrow e : \tau\|$$
$$= \|\varnothing\ \varnothing \vdash^{\hat{\Delta}; \hat{\Gamma}; \hat{\Psi}, \hat{a} \rightsquigarrow a \hookrightarrow \mathtt{uetsm}(\tau; e_{\mathrm{parse}}); b} \grave{e} \rightsquigarrow e : \tau\|$$
$$= \|b\|$$

The only case in the proof of part 2 that invokes part 1 is Case (3.6l). There, we have that $\mathsf{parseUExp}(\mathsf{subseq}(b; m; n)) = \hat{e}$ and the IH is applied to the judgement $\hat{\Delta}_{\mathrm{app}}\ \hat{\Gamma}_{\mathrm{app}} \vdash_{\hat{\Psi}} \hat{e} \rightsquigarrow e : \tau$ where $\hat{\Delta}_{\mathrm{app}} = \langle \mathcal{D}; \Delta_{\mathrm{app}} \rangle$ and $\hat{\Gamma}_{\mathrm{app}} = \langle \mathcal{G}; \Gamma_{\mathrm{app}} \rangle$ and $\hat{\Psi} = \langle \mathcal{A}; \Psi \rangle$. Because the metric is stable when passing from part 1 to part 2, we must have that it is strictly decreasing in the other direction:

$$\|\hat{\Delta}_{\mathrm{app}}\ \hat{\Gamma}_{\mathrm{app}} \vdash_{\hat{\Psi}} \hat{e} \rightsquigarrow e : \tau\| < \|\Delta\ \Gamma \vdash^{\hat{\Delta}_{\mathrm{app}}; \hat{\Gamma}_{\mathrm{app}}; \hat{\Psi}; b} \mathtt{cesplicede}[m; n] \rightsquigarrow e : \tau\|$$

i.e. by the definitions above,

$$\|\hat{e}\| < \|b\|$$

This is established by appeal to the following two conditions. The first condition simply states that subsequences of b are no longer than b.

Condition 3.22 (Body Subsequencing). *If* $\mathsf{subseq}(b; m; n) = b'$ *then* $\|b'\| \leq \|b\|$.

The second condition states that an unexpanded expression constructed by parsing a textual sequence b is strictly smaller, as measured by the metric defined above, than the length of b, because some characters must necessarily be used to invoke a TSM and delimit each literal body.

Condition 3.23 (Expression Parsing Monotonicity). *If* $\mathsf{parseUExp}(b) = \hat{e}$ *then* $\|\hat{e}\| < \|b\|$.

Combining Conditions 3.22 and 3.23, we have that $\|\hat{e}\| < \|b\|$ as needed. $\square$

Chapter 4

Unparameterized Pattern TSMs (upTSMs)

In Chapter 3, we considered situations where the programmer needed to *construct* (a.k.a. *introduce*) a value. In this chapter, we consider situations where the programmer needs to *deconstruct* (a.k.a. *eliminate*) a value. In full-scale functional languages like ML and Haskell, values are deconstructed by *pattern matching* over their structure. For example, recall the recursive labeled sum type `Rx` defined in Figure 2.1. We can pattern match over a value, r, of type `Rx` using VerseML's **match** construct:

```
1  fun read_example_rx(r : Rx) : (string * Rx) option =>
2    match r with
3      Seq(Str(name), Seq(Str ": ", ssn)) => Some (name, ssn)
4    | _ => None
```

Match expressions consist of a *scrutinee*, here r, and a sequence of *rules* separated by vertical bars, |, in the concrete syntax. Each rule consists of a *pattern* and an expression called the corresponding *branch*, separated by a double arrow, =>, in the concrete syntax. When the match expression is evaluated, the value of the scrutinee is matched against each pattern sequentially. If the value matches, evaluation takes the corresponding branch. Variables in patterns match any value of the appropriate type. In the corresponding branch, the variable stands for that value. Variables can each appear only once in a pattern. For example, on Line 3, the pattern `Seq(Str(name), Seq(Str ": ", ssn))` matches values of the form `Seq(Str(`e_1`), Seq(Str ": ", `e_2`))`, where e_1 is a value of type `string` and e_2 is a value of type `Rx`. The variables `name` and `ssn` stand for the values of e_1 and e_2, respectively, in `Some (name, ssn)`. On Line 4, the pattern `_` is the *wildcard pattern* – it matches any value of the appropriate type and binds no variables.

The behavior of the **match** construct when no pattern in the rule sequence matches a value is to raise an exception indicating *match failure*. It is possible to statically determine whether match failure is possible (i.e. whether there exist values of the scrutinee that are not matched by any pattern in the rule sequence). In the example above, our use of the wildcard pattern ensures that match failure cannot occur. A rule sequence that cannot lead to match failure is said to be *exhaustive*. Most compilers warn the programmer when a rule sequence is non-exhaustive.

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It is also possible to statically decide when a rule is *redundant* relative to the preceding rules, i.e. when there does not exist a value matched by that rule but not matched by any of the preceding rules. For example, if we add another rule at the end of the match expression above, it will be redundant because all values match the wildcard pattern. Again, most compilers warn the programmer when a rule is redundant.

Nested pattern matching generalizes the projection and case analysis operators (i.e. the *eliminators*) for products and sums (cf. miniVerse$_{UE}$ from the previous section) and decreases syntactic cost in situations where eliminators would need to be nested. There remains room for improvement, however, because complex patterns sometimes individually have high syntactic cost. In Sec. 2.3.2, we considered a hypothetical dialect of ML called ML+Rx that built in derived syntax both for constructing and pattern matching over values of the recursive labeled sum type Rx. In ML+Rx, we can express the example above at lower syntactic cost as follows:

```
1  fun read_example_rx(r : Rx) : (string * Rx) option =>
2    match r with
3      /@name: %ssn/ => Some (name, ssn)
4    | _ => None
```

Dialect formation is not a modular approach, for the reasons discussed in Chapter 1, so we seek language constructs that allow us to decrease the syntactic cost of expressing complex patterns to a similar degree.

Expression TSMs – introduced in Chapter 3 – can decrease the syntactic cost of constructing a value of a specified type. However, expressions are syntactically distinct from patterns, so we cannot simply apply an expression TSM to generate a pattern.[1] For this reason, we need to introduce a new (albeit closely related) construct – the **pattern TSM**. In this chapter, we consider only **unparameterized pattern TSMs** (upTSMs), i.e. pattern TSMs that generate patterns that match values of a single specified type, like Rx. In Chapter 5, we will consider both expression and pattern TSMs that specify type and module parameters (peTSMs and ppTSMs).

4.1   Pattern TSMs By Example

The organization of the remainder of this chapter mirrors that of Chapter 3. We begin in this section with a "tutorial-style" introduction to upTSMs in VerseML. In particular, we discuss an upTSM for patterns matching values of type Rx. In the next section, we specify a reduced formal system based on miniVerse$_{UE}$ called miniVerse$_{U}$ that makes the intuitions developed here mathematically precise.

---

[1]The fact that certain concrete expression and pattern forms overlap is immaterial to this fundamental distinction. There are many expression forms that the expansion generated by an expression TSM might use that have no corresponding pattern form, e.g. lambda abstraction.

4.1.1 Usage

The VerseML function `read_example_rx` defined at the beginning of this chapter can be concretely expressed at lower syntactic cost by applying a upTSM, `$rx`, as follows:

```
1  fun read_example_rx(r : Rx) : (string * Rx) option =>
2    match r with
3      $rx /@name: %ssn/ => Some (name, ssn)
4    | _ => None
```

Like expression TSMs, pattern TSMs are applied to *generalized literal forms* (see Figure 3.1). Generalized literal forms are left unparsed when patterns are first parsed. During the subsequent *typed expansion* process, the pattern TSM parses the body of the literal form to generate a *candidate expansion*. The language validates the candidate expansion according to criteria that we will establish in Sec. 4.1.4. If validation succeeds, the language generates the final expansion (or more concisely, simply the expansion) of the pattern. The expansion of the unexpanded pattern `$rx /@name: %ssn/` from the example above is the following pattern:

```
Seq(Str(name), Seq(Str ": ", ssn))
```

The checks for exhaustiveness and redundancy can be performed post-expansion in the usual way, so we do not need to consider them further here.

4.1.2 Definition

The definition of the pattern TSM `$rx` shown being applied in the example above has the following form:

```
syntax $rx at Rx for patterns {
  static fn(body : Body) : CEPat ParseResult =>
    (* regex pattern parser here *)
}
```

This definition first names the pattern TSM. Pattern TSM names, like expression TSM names, must begin with the dollar symbol ($) to distinguish them from labels. Pattern TSM names and expression TSM names are tracked separately, i.e. an expression TSM and a pattern TSM can have the same name without conflict (as is the case here – the expression TSM described in Sec. 3.1.2 is also named `$rx`). The *sort qualifier* **for patterns** indicates that this is a pattern TSM definition, rather than an expression TSM definition (the sort qualifier **for expressions** can be written for expression TSMs, though when the sort qualifier is omitted this is the default). Because defining both an expression TSM and a pattern TSM with the same name at the same type is a common idiom, VerseML provides a primitive derived form for combining their definitions:

```
syntax $rx at Rx for expressions {
  static fn(body : Body) : CEExp ParseResult =>
    (* regex expression parser here *)
} for patterns {
  static fn(body : Body) : CEPat ParseResult =>
    (* regex pattern parser here *)
```

```
type CEPat = Wild
           | (* ... *)
           | Spliced of IndexRange
```

Figure 4.1: Abbreviated definition of CEPat in the VerseML prelude.

}

Pattern TSMs, like expression TSMs, must specify a static *parse function*, delimited by curly braces in the concrete syntax. For a pattern TSM, the parse function must be of type Body -> CEPat ParseResult. The input type, Body, gives the parse function access to the body of the provided literal form, and is defined as in Sec. 3.1.2 as a synonym for the type string. The output type, CEPat ParseResult, is the parameterized type constructor ParseResult, defined in Figure 3.2, applied to the type CEPat defined in Figure 4.1. So if parsing succeeds, the pattern TSM returns a value of the form Success e_{cand} where e_{cand} is a value of type CEPat that we call the *encoding of the candidate expansion*. If parsing fails, then the pattern TSM returns a value constructed by ParseError and equipped with an error message and error location.

The type CEPat is analogous to the types CEExp and CETyp defined in Figure 3.3. It encodes the abstract syntax of VerseML patterns (in Figure 4.1, some constructors are elided for concision), with the exception of variable patterns (for reasons explained in Sec. 4.1.5 below), and includes an additional constructor, Spliced, for referring to spliced subpatterns by their position within the parse stream, discussed next.

4.1.3 Splicing

Patterns that appear directly within the literal body of an unexpanded pattern are called *spliced subpatterns*. For example, the patterns name and ssn appear within the unexpanded pattern $rx /@name: %ssn/. When the parse function determines that a subsequence of the literal body should be treated as a spliced subpatern (here, by recognizing the characters @ or % followed by a variable or parenthesized pattern), it can refer to it within the candidate expansion that it construct a reference to it for use within the candidate expansion it generates using the Spliced constructor of the CEPat type shown in Figure 4.1. The Spliced constructor requires a value of type IndexRange because spliced subpatterns are referred to indirectly by their position within the literal body. This prevents pattern TSMs from "forging" a spliced subpattern (i.e. claiming that some pattern is a spliced subpattern, even though it does not appear in the literal body).

The candidate expansion generated by the pattern TSM $rx for the example above, if written in a hypothetical concrete syntax where references to spliced subpatterns are written spliced<startIdx, endIdx>, is:

```
  Seq(Str(spliced<1, 4>), Seq(Str ": ", spliced<8, 10>))
```

Here, spliced<1, 4> refers to the subpattern name by position, and spliced<8, 10> refers to the subpattern ssn by position.

4.1.4 Typing

The language validates candidate expansion before a final expansion is generated. One aspect of candidate expansion validation is checking the candidate expansion against the type annotation specified by the pattern TSM, e.g. the type `Rx` in the example above.

4.1.5 Hygiene

In order to check that the candidate expansion is well-typed, the language must parse, type and expand the spliced subpatterns that the candidate expansion refers to (by their position within the literal body, cf. above). To maintain a useful binding discipline, i.e. to allow programmers to reason about variable binding without examining expansions directly, the validation process allows variables (e.g. `name` and `ssn` above) to occur only in spliced subpatterns (just as variables bound at the use site can only appear in spliced subexpressions when using TSMs). Indeed, there is no constructor for the type `CEPat` corresponding to a variable pattern. This protection against "hidden bindings" is beneficial because it leaves variable naming entirely up to the client of the pattern TSM. A pattern TSM cannot inadvertently shadow a binding at the application site.

4.1.6 Final Expansion

If validation succeeds, the semantics generates the *final expansion* of the pattern from the candidate expansion by replacing the references to spliced subpatterns with their final expansions. For example, the final expansion of `$rx /@name: %ssn/` is:

```
Seq(Str(name), Seq(Str ": ", ssn))
```

4.2 miniVerse$_\text{U}$

To make the intuitions developed in the previous section about pattern TSMs precise, we now introduce miniVerse$_\text{U}$, a reduced language with support for both ueTSMs and upTSMs. Like miniVerse$_\text{UE}$, miniVerse$_\text{U}$ consists of an *inner core* and an *outer surface*.

4.2.1 Syntax of the Inner Core

The *inner core* of miniVerse$_\text{U}$ consists of *types*, τ, *expanded expressions*, e, *expanded rules*, r, and *expanded patterns*, p. Their syntax is specified by the syntax chart in Figure 4.2. The inner core of miniVerse$_\text{U}$ forms a pure language and differs from the inner core of miniVerse$_\text{UE}$ only in that the case analysis operator has been replaced by the pattern matching operator[2], so we will gloss some definitions that would be expressed identically to those in Sec. 3.2. The new constructs are highlighted in gray in Figure 4.2. Our

---

[2]We retain the projection operator because it has lower syntactic cost than pattern matching when only a single field from a labeled tuple is needed.

formulation of the semantics of pattern matching is based on Harper's formulation in *Practical Foundations for Programming Languages, First Edition* [28].[3]

4.2.2 Statics of the Inner Core

The *statics of the inner core* is specified by judgements of the following form:

Judgement Form	Description
$\Delta \vdash \tau$ type	τ is a well-formed type assuming Δ
$\Delta\,\Gamma \vdash e : \tau$	e is assigned type τ assuming Δ and Γ
$\Delta\,\Gamma \vdash r : \tau \mapsto \tau'$	r takes values of type τ to values of type τ' assuming Δ and Γ
$\Delta \vdash p : \tau \dashv Y$	p matches values of type τ and generates hypotheses Y assuming Δ

The types of miniVerse$_U$ are exactly those of miniVerse$_{UE}$, described in Sec. 3.2, so the *type formation judgement*, $\Delta \vdash \tau$ type, is inductively defined by Rules (3.1).

The *typing judgement*, $\Delta\,\Gamma \vdash e : \tau$, assigns types to expressions and is inductively defined by Rules (4.1), which consist of:

- Rules written identically to Rules (3.2a) through (3.2j). We will refer to these rules as Rules (4.1a) through (4.1j).

- The following rule for match expressions:

$$\frac{\Delta\,\Gamma \vdash e : \tau \qquad \Delta \vdash \tau'\ \text{type} \qquad \{\Delta\,\Gamma \vdash r_i : \tau \mapsto \tau'\}_{1 \leq i \leq n}}{\Delta\,\Gamma \vdash \mathtt{match}[n]\{\tau'\}(e; \{r_i\}_{1 \leq i \leq n}) : \tau'} \tag{4.1k}$$

The first premise of Rule (4.1k) assigns a type, τ, to the scrutinee, e. The second premise checks that the type of the expression as a whole, τ', is well-formed.[4] The third premise then ensures that each rule r_i, for $1 \leq i \leq n$, takes values of type τ to values of the type of the match expression as a whole, τ'. This is expressed by the *rule typing judgement*, $\Delta\,\Gamma \vdash r : \tau \mapsto \tau'$, which is defined mutually with Rules (4.1) by the following rule:

$$\frac{\Delta \vdash p : \tau \dashv Y \qquad \Delta\,\Gamma \cup Y \vdash e : \tau'}{\Delta\,\Gamma \vdash \mathtt{rule}(p.e) : \tau \mapsto \tau'} \tag{4.2}$$

The premises of Rule (4.2) can be understood as follows, in order:

1. The first premise invokes the *pattern typing judgement*, $\Delta \vdash p : \tau \dashv Y$, to check that the pattern, p, matches values of type τ (defined assuming Δ), and to gather the typing hypotheses that the pattern generates in a *typing context*, Y. We use the metavariable Y (i.e. "upsilon") rather than Γ only to emphasize the distinct role of the typing context in the pattern typing judgement – algorithmically, it is the "output" of the judgement.

---

[3]The chapter on pattern matching has, of this writing, been removed from the draft second edition of *PFPL*, but a copy of the first edition can be found online.

[4]The second premise of Rule (4.1k), and the type argument in the match form, are necessary to maintain regularity, defined below, but only because when $n = 0$, the type τ' is arbitrary. In all other cases, τ' can be determined by assigning types to the branch expressions.

Sort			Operational Form	Stylized Form	Description
Typ	τ	$::=$	t	t	variable
			$\texttt{parr}(\tau;\tau)$	$\tau \rightharpoonup \tau$	partial function
			$\texttt{all}(t.\tau)$	$\forall t.\tau$	polymorphic
			$\texttt{rec}(t.\tau)$	$\mu t.\tau$	recursive
			$\texttt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L})$	$\langle \{i \hookrightarrow \tau_i\}_{i \in L} \rangle$	labeled product
			$\texttt{sum}[L](\{i \hookrightarrow \tau_i\}_{i \in L})$	$[\{i \hookrightarrow \tau_i\}_{i \in L}]$	labeled sum
Exp	e	$::=$	x	x	variable
			$\texttt{lam}\{\tau\}(x.e)$	$\lambda x{:}\tau.e$	abstraction
			$\texttt{ap}(e;e)$	$e(e)$	application
			$\texttt{tlam}(t.e)$	$\Lambda t.e$	type abstraction
			$\texttt{tap}\{\tau\}(e)$	$e[\tau]$	type application
			$\texttt{fold}\{t.\tau\}(e)$	$\texttt{fold}(e)$	fold
			$\texttt{unfold}(e)$	$\texttt{unfold}(e)$	unfold
			$\texttt{tpl}[L](\{i \hookrightarrow e_i\}_{i \in L})$	$\langle \{i \hookrightarrow e_i\}_{i \in L} \rangle$	labeled tuple
			$\texttt{pr}[\ell](e)$	$e \cdot \ell$	projection
			$\texttt{in}[L][\ell]\{\{i \hookrightarrow \tau_i\}_{i \in L}\}(e)$	$\texttt{inj}\ \ell \cdot e$	injection
			$\texttt{match}[n]\{\tau\}(e; \{r_i\}_{1 \le i \le n})$	$\texttt{match}\ e\ \{r_i\}_{1 \le i \le n}$	match
Rule	r	$::=$	$\texttt{rule}(p.e)$	$p \Rightarrow e$	rule
Pat	p	$::=$	x	x	variable pattern
			$\texttt{wildp}$	$\_$	wildcard pattern
			$\texttt{foldp}(p)$	$\texttt{fold}(p)$	fold pattern
			$\texttt{tplp}[L](\{i \hookrightarrow p_i\}_{i \in L})$	$\langle \{i \hookrightarrow p_i\}_{i \in L} \rangle$	labeled tuple pattern
			$\texttt{inp}[\ell](p)$	$\texttt{inj}\ \ell \cdot p$	injection pattern

Figure 4.2: Syntax of types and expanded expressions, rules and patterns (collectively, expanded terms) in miniVerse$_U$. We adopt the metatheoretic conventions established for our specification of miniVerse$_{UE}$ in Sec. 3.2 without restating them. We write $\{r_i\}_{1 \le i \le n}$ for sequences of $n \ge 0$ rule arguments and $p.e$ for expressions binding the variables that appear in the pattern p. Types and expanded terms are identified up to α-equivalence.

The pattern typing judgement is inductively defined by the following rules:

$$\frac{}{\Delta \vdash x : \tau \dashv\mid x : \tau} \tag{4.3a}$$

$$\frac{}{\Delta \vdash \texttt{wildp} : \tau \dashv\mid \varnothing} \tag{4.3b}$$

$$\frac{\Delta \vdash p : [\texttt{rec}(t.\tau)/t]\tau \dashv\mid Y}{\Delta \vdash \texttt{foldp}(p) : \texttt{rec}(t.\tau) \dashv\mid Y} \tag{4.3c}$$

$$\frac{\{\Delta \vdash p_i : \tau_i \dashv\mid Y_i\}_{i \in L}}{\Delta \vdash \texttt{tplp}[L](\{i \hookrightarrow p_i\}_{i \in L}) : \texttt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L}) \dashv\mid \cup_{i \in L} Y_i} \tag{4.3d}$$

$$\frac{\Delta \vdash p : \tau \dashv\mid Y}{\Delta \vdash \texttt{inp}[\ell](p) : \texttt{sum}[L, \ell](\{i \hookrightarrow \tau_i\}_{i \in L}; \ell \hookrightarrow \tau) \dashv\mid Y} \tag{4.3e}$$

Rule (4.3a) specifies that a variable pattern, x, matches values of any type, τ, and generates the hypothesis that x has the type τ.

Rule (4.3b) specifies that a wildcard pattern also matches values of any type, τ, but wildcard patterns generate no hypotheses.

Rule (4.3c) specifies that a fold pattern, $\texttt{foldp}(p)$, matches values of the recursive type $\texttt{rec}(t.\tau)$ and generates hypotheses Y if p matches values of a single unrolling of the recursive type, $[\texttt{rec}(t.\tau)/t]\tau$, and generates hypotheses Y.

Rule (4.3d) specifies that a labeled tuple pattern matches values of the labeled product type $\texttt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L})$. Labeled tuple patterns, $\texttt{tplp}[L](\{i \hookrightarrow p_i\}_{i \in L})$, specify a subpattern p_i for each label $i \in L$. The premise checks each subpattern p_i against the corresponding type τ_i, generating hypotheses Y_i. The conclusion of the rule gathers these hypotheses into a single pattern typing context, $\cup_{i \in L} Y_i$. The definition of typing context extension, applied iteratively here, implicitly requires that the pattern typing contexts Y_i be mutually disjoint, i.e.

$$\{\{\text{dom}(Y_i) \cap \text{dom}(Y_j) = \varnothing\}_{j \in L \setminus i}\}_{i \in L}$$

Rule (4.3e) specifies that an injection pattern, $\texttt{inp}[\ell](p)$, matches values of labeled sum types of the form $\texttt{sum}[L, \ell](\{i \hookrightarrow \tau_i\}_{i \in L}; \ell \hookrightarrow \tau)$, i.e. labeled sum types that define a case for the label ℓ, generating hypotheses Y if p matches value of type τ and generates hypotheses Y.

2. The final premise of Rule (4.2) extends the typing context, Γ, with the hypotheses generated by pattern typing, Y, and checks the branch expression, e, against the branch type, τ'.

The rules above are syntax-directed, so we assume an inversion lemma for each rule as needed without stating it separately or proving it explicitly. The following standard lemmas also hold.

The Weakening Lemma establishes that extending the context with unnecessary hypotheses preserves well-formedness and typing.

Lemma 4.1 (Weakening). *All of the following hold:*

1. *If $\Delta \vdash \tau$ type then Δ, t type $\vdash \tau$ type.*
2. (a) *If $\Delta\, \Gamma \vdash e : \tau$ then Δ, t type $\Gamma \vdash e : \tau$.*
   (b) *If $\Delta\, \Gamma \vdash r : \tau \Mapsto \tau'$ then Δ, t type $\Gamma \vdash r : \tau \Mapsto \tau'$.*
3. (a) *If $\Delta\, \Gamma \vdash e : \tau$ and $\Delta \vdash \tau''$ type then $\Delta\, \Gamma, x : \tau'' \vdash e : \tau$.*
   (b) *If $\Delta\, \Gamma \vdash r : \tau \Mapsto \tau'$ and $\Delta \vdash \tau''$ type then $\Delta\, \Gamma, x : \tau'' \vdash r : \tau \Mapsto \tau'$.*
4. *If $\Delta \vdash p : \tau \dashv Y$ then Δ, t type $\vdash p : \tau \dashv Y$.*

Proof Sketch.

1. By rule induction over Rules (3.1).
2. By mutual rule induction over Rules (4.1) and Rule (4.2) and part 1.
3. By mutual rule induction over Rules (4.1) and Rule (4.2) and part 1.
4. By rule induction over Rules (4.3).

$\square$

The pattern typing judgement is a *linear* in the pattern typing context, i.e. it does *not* obey weakening of the pattern typing context. This is to ensure that the pattern typing context captures exactly those hypotheses generated by a pattern, and no others.

We assume that renaming of bound identifiers, α-equivalence and substitution can be defined essentially as in *PFPL* [29], modified only so that binders involving patterns bind exactly those variables mentioned in the pattern in some arbitrary deterministic order. The Substitution Lemma establishes that substitution of a well-formed type for a type variable, or an expanded expression of the appropriate type for an expanded expression variable, preserves well-formedness and typing.

Lemma 4.2 (Substitution). *All of the following hold:*

1. *If Δ, t type $\vdash \tau$ type and $\Delta \vdash \tau'$ type then $\Delta \vdash [\tau'/t]\tau$ type.*
2. (a) *If Δ, t type $\Gamma \vdash e : \tau$ and $\Delta \vdash \tau'$ type then $\Delta\, [\tau'/t]\Gamma \vdash [\tau'/t]e : [\tau'/t]\tau$.*
   (b) *If Δ, t type $\Gamma \vdash r : \tau \Mapsto \tau''$ and $\Delta \vdash \tau'$ type then $\Delta\, [\tau'/t]\Gamma \vdash [\tau'/t]r : [\tau'/t]\tau \Mapsto [\tau'/t]\tau''$.*
3. (a) *If $\Delta\, \Gamma, x : \tau' \vdash e : \tau$ and $\Delta\, \Gamma \vdash e' : \tau'$ then $\Delta\, \Gamma \vdash [e'/x]e : \tau$.*
   (b) *If $\Delta\, \Gamma, x : \tau' \vdash r : \tau \Mapsto \tau''$ and $\Delta\, \Gamma \vdash e' : \tau''$ then $\Delta\, \Gamma \vdash [e'/x]r : \tau \Mapsto \tau''$.*

Proof Sketch.

1. By rule induction over Rules (3.1).
2. By mutual rule induction over Rules (4.1) and Rule (4.2).
3. By mutual rule induction over Rules (4.1) and Rule (4.2).

$\square$

The Decomposition Lemma is the converse of the Substitution Lemma.

Lemma 4.3 (Decomposition). *All of the following hold:*

1. *If $\Delta \vdash [\tau'/t]\tau$ type and $\Delta \vdash \tau'$ type then Δ, t type $\vdash \tau$ type.*
2. (a) *If $\Delta\, [\tau'/t]\Gamma \vdash [\tau'/t]e : [\tau'/t]\tau$ and $\Delta \vdash \tau'$ type then Δ, t type $\Gamma \vdash e : \tau$.*
   (b) *If $\Delta\, [\tau'/t]\Gamma \vdash [\tau'/t]r : [\tau'/t]\tau \Mapsto [\tau'/t]\tau''$ and $\Delta \vdash \tau'$ type then Δ, t type $\Gamma \vdash r : \tau \Mapsto \tau''$.*
3. (a) *If $\Delta\, \Gamma \vdash [e'/x]e : \tau$ and $\Delta\, \Gamma \vdash e' : \tau'$ then $\Delta\, \Gamma, x : \tau' \vdash e : \tau$.*

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(b) If $\Delta\ \Gamma \vdash [e'/x]r : \tau \mapsto \tau''$ *and* $\Delta\ \Gamma \vdash e' : \tau'$ *then* $\Delta\ \Gamma, x : \tau' \vdash r : \tau \mapsto \tau''$.

Proof Sketch.

1. By rule induction over Rules (3.1) and case analysis over the definition of substitution. In all cases, the derivation of $\Delta \vdash [\tau'/t]\tau$ type does not depend on the form of τ'.

2. By mutual rule induction over Rules (4.1) and Rule (4.2) and case analysis over the definition of substitution. In all cases, the derivation of $\Delta\ [\tau'/t]\Gamma \vdash [\tau'/t]e : [\tau'/t]\tau$ or $\Delta\ [\tau'/t]\Gamma \vdash [\tau'/t]r : [\tau'/t]\tau \mapsto [\tau'/t]\tau''$ does not depend on the form of τ'.

3. By mutual rule induction over Rules (4.1) and Rule (4.2) and case analysis over the definition of substitution. In all cases, the derivation of $\Delta\ \Gamma \vdash [e'/x]e : \tau$ or $\Delta\ \Gamma \vdash [e'/x]r : \tau \mapsto \tau''$ does not depend on the form of e'.

$\square$

The Pattern Regularity Lemma establishes that the hypotheses generated by checking a pattern against a well-formed type involve only well-formed types.

Lemma 4.4 (Pattern Regularity)**.** *If* $\Delta \vdash p : \tau \dashv\vdash Y$ *and* $\Delta \vdash \tau$ type *then* $\Delta \vdash Y$ ctx.

Proof. By rule induction over Rules (4.3).

Case (4.3a)**.** We have:

(1) $p = x$	by assumption
(2) $Y = x : \tau$	by assumption
(3) $\Delta \vdash \tau$ type	by assumption
(4) $\Delta \vdash x : \tau$ ctx	by Definition 3.1 on (3)

Case (4.3b)**.** We have:

(1) $Y = \varnothing$	by assumption
(2) $\Delta \vdash \varnothing$ ctx	by Definition 3.1

Case (4.3d)**.** We have:

(1) $p = \mathtt{tplp}[L](\{i \hookrightarrow p_i\}_{i\in L})$	by assumption
(2) $\tau = \mathtt{prod}[L](\{i \hookrightarrow \tau_i\}_{i\in L})$	by assumption
(3) $Y = \cup_{i\in L} Y_i$	by assumption
(4) $\{\Delta \vdash p_i : \tau_i \dashv\vdash Y_i\}_{i\in L}$	by assumption
(5) $\Delta \vdash \mathtt{prod}[L](\{i \hookrightarrow \tau_i\}_{i\in L})$ type	by assumption
(6) $\{\Delta \vdash \tau_i$ type$\}_{i\in L}$	by Inversion of Rule (3.1e) on (5)
(7) $\{\Delta \vdash Y_i$ ctx$\}_{i\in L}$	by IH over (4) and (6)

(8) $\Delta \vdash \cup_{i \in L} Y_i$ ctx

by Definition 3.1 on (7), then Definition 3.1 again, using the definition of typing context union iteratively

Case (4.3e). We have:

(1) $p = \mathtt{inp}[\ell](p')$ by assumption
(2) $\tau = \mathtt{sum}[L, \ell](\{i \hookrightarrow \tau_i\}_{i \in L}; \ell \hookrightarrow \tau')$ by assumption
(3) $\Delta \vdash \mathtt{sum}[L, \ell](\{i \hookrightarrow \tau_i\}_{i \in L}; \ell \hookrightarrow \tau')$ type   by assumption
(4) $\Delta \vdash p' : \tau' \dashv\mathrel{\mkern-5mu}\mathrel{\mkern-5mu} Y$ by assumption
(5) $\Delta \vdash \tau'$ type   by Inversion of Rule (3.1f) on (3)
(6) $\Delta \vdash Y$ ctx   by IH on (4) and (5)

$\square$

Finally, the Regularity Lemma establishes that the type assigned to an expression under a well-formed typing context is well-formed.

Lemma 4.5 (Regularity). *All of the following hold:*

1. *If $\Delta\,\Gamma \vdash e : \tau$ and $\Delta \vdash \Gamma$ ctx then $\Delta \vdash \tau$ type.*
2. *If $\Delta\,\Gamma \vdash r : \tau \mapsto \tau'$ and $\Delta \vdash \Gamma$ ctx then $\Delta \vdash \tau'$ type.*

Proof Sketch. By mutual rule induction over Rules (4.1) and Rule (4.2), and Lemma 4.2 and Lemma 4.4. $\square$

4.2.3 Structural Dynamics

The *structural dynamics of* miniVerse$_U$ is specified as a transition system, and is organized around judgements of the following form:

Judgement Form	Description
$e \mapsto e'$	e transitions to e'
e val	e is a value
e matchfail	e raises match failure

We also define auxiliary judgements for *iterated transition, $e \mapsto^* e'$,* and *evaluation, $e \Downarrow e'$.*

Definition 4.6 (Iterated Transition). *Iterated transition, $e \mapsto^* e'$, is the reflexive, transitive closure of the transition judgement, $e \mapsto e'$.*

Definition 4.7 (Evaluation). *$e \Downarrow e'$ iff $e \mapsto^* e'$ and e' val.*

As in Sec. 3.2.3, our subsequent developments do not make mention of particular rules in the dynamics, nor do they make mention of judgements that are used only for defining the dynamics of the match operator, so we do not produce these details here. Instead, it suffices to state the following conditions.

The Canonical Forms condition characterizes well-typed values. Satisfying this condition requires an *eager* (i.e. *by-value*) formulation of the dynamics. This condition is identical to Condition 3.8.

Condition 4.8 (Canonical Forms). *If $\vdash e : \tau$ and e val then:*

1. *If $\tau = \mathtt{parr}(\tau_1; \tau_2)$ then $e = \mathtt{lam}\{\tau_1\}(x.e')$ and $x : \tau_1 \vdash e' : \tau_2$.*
2. *If $\tau = \mathtt{all}(t.\tau')$ then $e = \mathtt{tlam}(t.e')$ and t type $\vdash e' : \tau'$.*
3. *If $\tau = \mathtt{rec}(t.\tau')$ then $e = \mathtt{fold}\{t.\tau'\}(e')$ and $\vdash e' : [\mathtt{rec}(t.\tau')/t]\tau'$ and e' val.*
4. *If $\tau = \mathtt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L})$ then $e = \mathtt{tpl}[L](\{i \hookrightarrow e_i\}_{i \in L})$ and $\vdash e_i : \tau_i$ and e_i val for each $i \in L$.*
5. *If $\tau = \mathtt{sum}[L](\{i \hookrightarrow \tau_i\}_{i \in L})$ then for some label set L' and label ℓ and type τ_ℓ, we have that $L = L', \ell$ and $\tau = \mathtt{sum}[L', \ell](\{i \hookrightarrow \tau_i\}_{i \in L'}; \ell \hookrightarrow \tau_\ell)$ and $e = \mathtt{in}[L', \ell][\ell]\{\{i \hookrightarrow \tau_i\}_{i \in L'}; \ell \hookrightarrow \tau_\ell\}(e')$ and $\vdash e' : \tau_\ell$ and e' val.*

The Preservation condition ensures that evaluation preserves typing.

Condition 4.9 (Preservation). *If $\vdash e : \tau$ and $e \mapsto e'$ then $\vdash e' : \tau$.*

The Progress condition ensures that evaluation of a well-typed expanded expression cannot "get stuck".

Condition 4.10 (Progress). *If $\vdash e : \tau$ then either e val or e matchfail or there exists an e' such that $e \mapsto e'$.*

Together, these two conditions constitute the Type Safety Condition.

We do not define exhaustiveness and redundancy properties here, because these can be checked post-expansion and so are also not relevant to our subsequent developments (but see [28] for a formal account).

4.2.4   Syntax of the Outer Surface

A miniVerse$_\mathbf{U}$ program ultimately evaluates as an expanded expression. However, the programmer does not write the expanded expression directly. Instead, the programmer writes a textual sequence, b, consisting of characters in some suitable alphabet (e.g. in practice, `ASCII` or `Unicode`), which is parsed by some partial metafunction parseUExp(b) to produce an *unexpanded expression*, $\hat{e}$. Unexpanded expressions can contain *unexpanded types*, $\hat{\tau}$, *unexpanded rules*, $\hat{r}$, and *unexpanded patterns*, $\hat{p}$, so we also need partial metafunctions parseUTyp(b), parseURule(b) and parseUPat(b). The abstract syntax of unexpanded types, expressions, rules and patterns, which form the *outer surface* of miniVerse$_\mathbf{U}$, is defined in Figure 4.3. The full definition of the textual syntax of miniVerse$_\mathbf{U}$ is not important for our purposes, so we simply give the following condition, which states that there is some way to textually represent every unexpanded type, expression, rule and pattern.

Condition 4.11 (Textual Representability). *All of the following must hold:*

1. *For each $\hat{\tau}$, there exists b such that parseUTyp(b) = $\hat{\tau}$.*
2. *For each $\hat{e}$, there exists b such that parseUExp(b) = $\hat{e}$.*
3. *For each $\hat{r}$, there exists b such that parseURule(b) = $\hat{r}$.*
4. *For each $\hat{p}$, there exists b such that parseUPat(b) = $\hat{p}$.*

Sort		Operational Form	Stylized Form	Description
UTyp	$\hat{\tau}$ ::=	$\hat{t}$	$\hat{t}$	sigil
		$\mathtt{uparr}(\hat{\tau};\hat{\tau})$	$\hat{\tau} \rightharpoonup \hat{\tau}$	partial function
		$\mathtt{uall}(\hat{t}.\hat{\tau})$	$\forall \hat{t}.\hat{\tau}$	polymorphic
		$\mathtt{urec}(\hat{t}.\hat{\tau})$	$\mu\hat{t}.\hat{\tau}$	recursive
		$\mathtt{uprod}[L](\{i \hookrightarrow \hat{\tau}_i\}_{i\in L})$	$\langle \{i \hookrightarrow \hat{\tau}_i\}_{i\in L}\rangle$	labeled product
		$\mathtt{usum}[L](\{i \hookrightarrow \hat{\tau}_i\}_{i\in L})$	$[\{i \hookrightarrow \hat{\tau}_i\}_{i\in L}]$	labeled sum
UExp	$\hat{e}$ ::=	$\hat{x}$	$\hat{x}$	sigil
		$\mathtt{ulam}\{\hat{\tau}\}(\hat{x}.\hat{e})$	$\lambda\hat{x}{:}\hat{\tau}.\hat{e}$	abstraction
		$\mathtt{uap}(\hat{e};\hat{e})$	$\hat{e}(\hat{e})$	application
		$\mathtt{utlam}(\hat{t}.\hat{e})$	$\Lambda\hat{t}.\hat{e}$	type abstraction
		$\mathtt{utap}\{\hat{\tau}\}(\hat{e})$	$\hat{e}[\hat{\tau}]$	type application
		$\mathtt{ufold}\{\hat{t}.\hat{\tau}\}(\hat{e})$	$\mathtt{fold}(\hat{e})$	fold
		$\mathtt{uunfold}(\hat{e})$	$\mathtt{unfold}(\hat{e})$	unfold
		$\mathtt{utpl}[L](\{i \hookrightarrow \hat{e}_i\}_{i\in L})$	$\langle \{i \hookrightarrow \hat{e}_i\}_{i\in L}\rangle$	labeled tuple
		$\mathtt{upr}[\ell](\hat{e})$	$\hat{e}\cdot\ell$	projection
		$\mathtt{uin}[L][\ell]\{\{i \hookrightarrow \hat{\tau}_i\}_{i\in L}\}(\hat{e})$	$\mathtt{inj}\ \ell\cdot\hat{e}$	injection
		$\mathtt{umatch}[n]\{\hat{\tau}\}(\hat{e};\{\hat{r}_i\}_{1\le i\le n})$	$\mathtt{match}\ \hat{e}\ \{\hat{r}_i\}_{1\le i\le n}$	match
		$\mathtt{usyntaxue}\{e\}\{\hat{\tau}\}(\hat{a}.\hat{e})$	$\mathtt{syntax}\ \hat{a}\ \mathtt{at}\ \hat{\tau}\ \mathtt{for}$ $\mathtt{expressions}\ \{e\}\ \mathtt{in}\ \hat{e}$	ueTSM definition
		$\mathtt{uapuetsm}[b][\hat{a}]$	$\hat{a}\ /b/$	ueTSM application
		$\mathtt{usyntaxup}\{e\}\{\hat{\tau}\}(\hat{a}.\hat{e})$	$\mathtt{syntax}\ \hat{a}\ \mathtt{at}\ \hat{\tau}\ \mathtt{for}$ $\mathtt{patterns}\ \{e\}\ \mathtt{in}\ \hat{e}$	upTSM definition
URule	$\hat{r}$ ::=	$\mathtt{urule}(\hat{p}.\hat{e})$	$\hat{p} \Rightarrow \hat{e}$	match rule
UPat	$\hat{p}$ ::=	$\hat{x}$	$\hat{x}$	sigil pattern
		$\mathtt{uwildp}$	$\_$	wildcard pattern
		$\mathtt{ufoldp}(\hat{p})$	$\mathtt{fold}(\hat{p})$	fold pattern
		$\mathtt{utplp}[L](\{i \hookrightarrow \hat{p}_i\}_{i\in L})$	$\langle \{i \hookrightarrow \hat{p}_i\}_{i\in L}\rangle$	labeled tuple pattern
		$\mathtt{uinp}[\ell](\hat{p})$	$\mathtt{inj}\ \ell\cdot\hat{p}$	injection pattern
		$\mathtt{uapuptsm}[b][\hat{a}]$	$\hat{a}\ /b/$	upTSM application

Figure 4.3: Abstract syntax of unexpanded types, expressions, rules and patterns in miniVerse$_\mathbf{U}$. Metavariable $\hat{t}$ ranges over type sigils, $\hat{x}$ ranges over expression sigils, $\hat{a}$ over TSM names and b over textual sequences, which, when they appear in an unexpanded term, are called literal bodies. Literal bodies might contain spliced subterms that are only "surfaced" during typed expansion, so renaming of bound identifiers and substitution are not defined over unexpanded types and terms.

As in miniVerse$_{\mathbf{UE}}$, unexpanded types and expressions bind *type sigils*, $\hat{t}$, *expression sigils*, $\hat{x}$, and *TSM names*, $\hat{a}$. Sigils are given meaning by expansion to variables during typed expansion. We **cannot** adopt the usual definition of α-renaming of identifiers, because unexpanded types and expressions are still in a "partially parsed" state – the literal bodies, b, within an unexpanded expression might contain spliced subterms that are "surfaced" by a TSM only during typed expansion, as we will detail below.

Each inner core form (defined in Figure 4.2) maps onto an outer surface form. We refer to these as the *shared forms*. In particular:

- Each type variable, t, maps onto a unique type sigil, written $\widehat{t}$ (pronounced "sigil of t"). Notice the distinction between $\hat{t}$, which is a metavariable ranging over type sigils, and $\widehat{t}$, which is a metafunction, written in stylized form, applied to a type variable to produce a type sigil.

- Each type form, τ, maps onto an unexpanded type form, $\mathcal{U}(\tau)$, according to the definition of $\mathcal{U}(\tau)$ in Sec. 3.2.4.

- Each expression variable, x, maps onto a unique expression sigil, written $\widehat{x}$. Again, notice the distinction between $\hat{x}$ and $\widehat{x}$.

- Each expanded expression form, e, maps onto an unexpanded expression form $\mathcal{U}(e)$ as follows:

$$\mathcal{U}(x) = \widehat{x}$$
$$\mathcal{U}(\mathtt{lam}\{\tau\}(x.e)) = \mathtt{ulam}\{\mathcal{U}(\tau)\}(\widehat{x}.\mathcal{U}(e))$$
$$\mathcal{U}(\mathtt{ap}(e_1;e_2)) = \mathtt{uap}(\mathcal{U}(e_1);\mathcal{U}(e_2))$$
$$\mathcal{U}(\mathtt{tlam}(t.e)) = \mathtt{utlam}(\widehat{t}.\mathcal{U}(e))$$
$$\mathcal{U}(\mathtt{tap}\{\tau\}(e)) = \mathtt{utap}\{\mathcal{U}(\tau)\}(\mathcal{U}(e))$$
$$\mathcal{U}(\mathtt{fold}\{t.\tau\}(e)) = \mathtt{ufold}\{\widehat{t}.\mathcal{U}(\tau)\}(\mathcal{U}(e))$$
$$\mathcal{U}(\mathtt{unfold}(e)) = \mathtt{uunfold}(\mathcal{U}(e))$$
$$\mathcal{U}(\mathtt{tpl}[L](\{i \hookrightarrow e_i\}_{i \in L})) = \mathtt{utpl}[L](\{i \hookrightarrow \mathcal{U}(e_i)\}_{i \in L})$$
$$\mathcal{U}(\mathtt{in}[L][\ell]\{\{i \hookrightarrow \tau_i\}_{i \in L}\}(e)) = \mathtt{uin}[L][\ell]\{\{i \hookrightarrow \mathcal{U}(\tau_i)\}_{i \in L}\}(\mathcal{U}(e))$$
$$\mathcal{U}(\mathtt{match}[n]\{\tau\}(e;\{r_i\}_{1 \leq i \leq n})) = \mathtt{umatch}[n]\{\mathcal{U}(\tau)\}(\mathcal{U}(e);\{\mathcal{U}(r_i)\}_{1 \leq i \leq n})$$

- The expanded rule form maps onto the unexpanded rule form as follows:

$$\mathcal{U}(\mathtt{rule}(p.e)) = \mathtt{urule}(\mathcal{U}(p).\mathcal{U}(e))$$

- Each expanded pattern form, p, maps onto the unexpanded pattern form $\mathcal{U}(p)$ as follows:

$$\mathcal{U}(x) = \widehat{x}$$
$$\mathcal{U}(\mathtt{wildp}) = \mathtt{uwildp}$$
$$\mathcal{U}(\mathtt{foldp}(p)) = \mathtt{ufoldp}(\mathcal{U}(p))$$

$$\mathcal{U}(\texttt{tplp}[L](\{i \hookrightarrow p_i\}_{i \in L})) = \texttt{utplp}[L](\{i \hookrightarrow \mathcal{U}(p_i)\}_{i \in L})$$
$$\mathcal{U}(\texttt{inp}[\ell](p)) = \texttt{uinp}[\ell](\mathcal{U}(p))$$

The only unexpanded forms that do not correspond to expanded forms are the unexpanded expression forms for ueTSM definition, ueTSM application and upTSM definition, and the unexpanded pattern form for upTSM application. The forms related to upTSMs are highlighted in gray in Figure 4.3.

4.2.5 Typed Expansion

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

Judgement Form	Description
$\hat{\Delta} \vdash \hat{\tau} \rightsquigarrow \tau$ type	$\hat{\tau}$ is well-formed and has expansion τ assuming $\hat{\Delta}$
$\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \rightsquigarrow e : \tau$	$\hat{e}$ has expansion e and type τ under ueTSM context $\hat{\Psi}$ and upTSM context $\hat{\Phi}$ assuming $\hat{\Delta}$ and $\hat{\Gamma}$
$\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{r} \rightsquigarrow r : \tau \Mapsto \tau'$	$\hat{r}$ has expansion r and takes values of type τ to values of type τ' under $\hat{\Psi}$ and $\hat{\Phi}$ assuming $\hat{\Delta}$ and $\hat{\Gamma}$
$\Delta \vdash_{\hat{\Phi}} \hat{p} \rightsquigarrow p : \tau \dashv \hat{Y}$	$\hat{p}$ has expansion p and type τ and generates hypotheses Y under upTSM context $\hat{\Phi}$ assuming Δ

Type Expansion

Unexpanded type formation contexts, $\hat{\Delta}$, were defined in Sec. 3.2.5. The *type expansion judgement*, $\hat{\Delta} \vdash \hat{\tau} \rightsquigarrow \tau$ type, is inductively defined by Rules (3.3).

Typed Expression, Rule and Pattern Expansion

Unexpanded typing contexts, $\hat{\Gamma}$, were defined in Sec. 3.2.5. *Unexpanded pattern typing contexts*, $\hat{Y}$, are defined identically to unexpanded typing contexts (i.e. we only use a distinct metavariable to emphasize their distinct roles in the judgements above).

The *typed expression expansion* judgement, $\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \rightsquigarrow e : \tau$, and the *typed rule expansion judgement*, $\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{r} \rightsquigarrow r : \tau \Mapsto \tau'$ are defined mutually inductively by Rules (4.4) and Rule (4.5), respectively, and the *typed pattern expansion judgement*, $\Delta \vdash_{\Phi} \hat{p} \rightsquigarrow p : \tau \dashv Y$, is inductively defined by Rules (4.6) as follows.

Shared Forms Rules (4.4a) through (4.4k) define typed expansion of unexpanded expressions of shared form. The first five of these rules are shown below:

$$\frac{}{\hat{\Delta}\,\hat{\Gamma}, \hat{x} \rightsquigarrow x : \tau \vdash_{\hat{\Psi};\hat{\Phi}} \hat{x} \rightsquigarrow x : \tau} \tag{4.4a}$$

$$\dfrac{\hat{\Delta} \vdash \hat{\tau} \rightsquigarrow \tau \text{ type} \qquad \hat{\Delta}\,\hat{\Gamma}, \hat{x} \rightsquigarrow x : \tau \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \rightsquigarrow e : \tau'}{\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \mathtt{ulam}\{\hat{\tau}\}(\hat{x}.\hat{e}) \rightsquigarrow \mathtt{lam}\{\tau\}(x.e) : \mathtt{parr}(\tau;\tau')} \tag{4.4b}$$

$$\dfrac{\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e}_1 \rightsquigarrow e_1 : \mathtt{parr}(\tau;\tau') \qquad \hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e}_2 \rightsquigarrow e_2 : \tau}{\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \mathtt{uap}(\hat{e}_1;\hat{e}_2) \rightsquigarrow \mathtt{ap}(e_1;e_2) : \tau'} \tag{4.4c}$$

$$\dfrac{\hat{\Delta}, \hat{t} \rightsquigarrow t \text{ type}\, \hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \rightsquigarrow e : \tau}{\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \mathtt{utlam}(\hat{t}.\hat{e}) \rightsquigarrow \mathtt{tlam}(t.e) : \mathtt{all}(t.\tau)} \tag{4.4d}$$

$$\dfrac{\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \rightsquigarrow e : \mathtt{all}(t.\tau) \qquad \hat{\Delta} \vdash \hat{\tau}' \rightsquigarrow \tau' \text{ type}}{\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \mathtt{utap}\{\hat{\tau}'\}(\hat{e}) \rightsquigarrow \mathtt{tap}\{\tau'\}(e) : [\tau'/t]\tau} \tag{4.4e}$$

These rules are similar to Rules (3.4a) through (3.4e). In particular, in both sets of rules, the unexpanded and expanded expression forms in the conclusion correspond, and the premises correspond to those of the typing rule for the expanded expression form – here, Rules (4.1a) through (4.1e), respectively. In particular, each type expansion premise in each rule above corresponds to a type formation premise in the corresponding typing rule, and each typed expression expansion premise in each rule above corresponds to a typing premise in the corresponding typing rule. The type assigned in the conclusion of each rule above is identical to the type assigned in the conclusion of the corresponding typing rule. The ueTSM context, $\hat{\Psi}$, and now also the upTSM context, $\hat{\Phi}$, pass opaquely through these rules.

Rule (4.4k), below, handles unexpanded match expressions and corresponds in the same way to Rule (4.1k).

$$\dfrac{\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \rightsquigarrow e : \tau \qquad \hat{\Delta} \vdash \hat{\tau}' \rightsquigarrow \tau' \text{ type} \qquad \{\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{r}_i \rightsquigarrow r_i : \tau \Mapsto \tau'\}_{1 \leq i \leq n}}{\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \mathtt{umatch}[n]\{\hat{\tau}'\}(\hat{e}; \{\hat{r}_i\}_{1 \leq i \leq n}) \rightsquigarrow \mathtt{match}[n]\{\tau'\}(e; \{r_i\}_{1 \leq i \leq n}) : \tau'} \tag{4.4k}$$

We can express this scheme more precisely with the following rule transformation. For each rule in Rules (4.1),

$$\dfrac{J_1 \qquad \cdots \qquad J_k}{J}$$

the corresponding typed expansion rule is

$$\dfrac{\mathcal{U}(J_1) \qquad \cdots \qquad \mathcal{U}(J_k)}{\mathcal{U}(J)}$$

where

$$\mathcal{U}(\Delta \vdash \tau \text{ type}) = \mathcal{U}(\Delta) \vdash \mathcal{U}(\tau) \rightsquigarrow \tau \text{ type}$$
$$\mathcal{U}(\Gamma\,\Delta \vdash e : \tau) = \mathcal{U}(\Gamma)\,\mathcal{U}(\Delta) \vdash_{\hat{\Psi};\hat{\Phi}} \mathcal{U}(e) \rightsquigarrow e : \tau$$
$$\mathcal{U}(\Gamma\,\Delta \vdash r : \tau \Mapsto \tau') = \mathcal{U}(\Gamma)\,\mathcal{U}(\Delta) \vdash_{\hat{\Psi};\hat{\Phi}} \mathcal{U}(r) \rightsquigarrow r : \tau \Mapsto \tau'$$
$$\mathcal{U}(\{J_i\}_{i \in L}) = \{\mathcal{U}(J_i)\}_{i \in L}$$

and where $\mathcal{U}(\Delta), \mathcal{U}(\Gamma)$ and $\mathcal{U}(\tau)$ are defined as in Sec. 3.2.5 and:

- $\mathcal{U}(e)$ is defined as follows
  - When e is of definite form, $\mathcal{U}(e)$ is defined as in Sec. 4.2.4.
  - When e is of indefinite form, $\mathcal{U}(e)$ is a uniquely corresponding metavariable of sort UExp also of indefinite form. For example, $\mathcal{U}(e_1) = \hat{e}_1$ and $\mathcal{U}(e_2) = \hat{e}_2$.
- $\mathcal{U}(r)$ is defined as follows:
  - When r is of definite form, $\mathcal{U}(r)$ is defined as in Sec. 4.2.4.
  - When e is of indefinite form, $\mathcal{U}(r)$ is a uniquely corresponding metavariable of sort URule also of indefinite form.

It is instructive to use this rule transformation to generate Rules (4.4a) through (4.4e) and Rule (4.4k) above. We omit the remaining rules generated by this transformation, i.e. Rules (4.4d) through (4.4j).

The typed rule expansion judgement is defined by Rule (4.5), below.

$$\frac{\Delta \vdash_{\hat{\Phi}} \hat{p} \rightsquigarrow p : \tau \dashv\!\vert \langle \mathcal{G}'; Y \rangle \qquad \langle \mathcal{D}; \Delta \rangle \, \langle \mathcal{G} \uplus \mathcal{G}'; \Gamma \cup Y \rangle \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{e} \rightsquigarrow e : \tau'}{\langle \mathcal{D}; \Delta \rangle \, \langle \mathcal{G}; \Gamma \rangle \vdash_{\hat{\Psi}; \hat{\Phi}} \mathtt{urule}(\hat{p}.\hat{e}) \rightsquigarrow \mathtt{rule}(p.e) : \tau \Rightarrow \tau'} \quad (4.5)$$

As in the typed expression expansion judgements, the unexpanded and expanded forms in the conclusion of the rule above correspond. The premises correspond to those of Rule (4.2). In particular, the typed pattern expansion premise in the rule above corresponds to the pattern typing premise of Rule (4.2) and the typed expression expansion premise in the rule above corresponds to the typing premise of Rule (4.2). Because unexpanded terms bind expression sigils, which are given meaning by expansion to variables, the pattern typing rules must generate both a sigil expansion context, $\mathcal{G}'$, and a pattern typing context, Y. In the second premise, we update the "incoming" sigil expansion context, $\mathcal{G}$, with the new sigil expansions, $\mathcal{G}'$, and correspondingly, extend the "incoming" typing context, Γ, with the new hypotheses, Y.

Rules (4.6a) through (4.6e), below, define typed expansion of unexpanded patterns of shared form.

$$\frac{}{\Delta \vdash_{\hat{\Phi}} \hat{x} \rightsquigarrow x : \tau \dashv\!\vert \langle \hat{x} \rightsquigarrow x; x : \tau \rangle} \quad (4.6a)$$

$$\frac{}{\Delta \vdash_{\hat{\Phi}} \mathtt{uwildp} \rightsquigarrow \mathtt{wildp} : \tau \dashv\!\vert \langle \varnothing; \varnothing \rangle} \quad (4.6b)$$

$$\frac{\Delta \vdash_{\hat{\Phi}} \hat{p} \rightsquigarrow p : [\mathtt{rec}(t.\tau)/t]\tau \dashv\!\vert \hat{Y}}{\Delta \vdash_{\hat{\Phi}} \mathtt{ufoldp}(\hat{p}) \rightsquigarrow \mathtt{foldp}(p) : \mathtt{rec}(t.\tau) \dashv\!\vert \hat{Y}} \quad (4.6c)$$

$$\left(\frac{\{\Delta \vdash_{\hat{\Phi}} \hat{p}_i \rightsquigarrow p_i : \tau_i \dashv\!\vert \hat{Y}_i\}_{i \in L}}{\begin{array}{c} \Delta \vdash_{\hat{\Phi}} \mathtt{utplp}[L](\{i \hookrightarrow \hat{p}_i\}_{i \in L}) \\ \rightsquigarrow \\ \mathtt{tplp}[L](\{i \hookrightarrow p_i\}_{i \in L}) : \mathtt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L}) \dashv\!\vert \cup_{i \in L} \hat{Y}_i \end{array}} \right) \quad (4.6d)$$

$$\frac{\Delta \vdash_{\hat{\Phi}} \hat{p} \rightsquigarrow p : \tau \dashv\!\vert \hat{Y}}{\Delta \vdash_{\hat{\Phi}} \mathtt{uinp}[\ell](\hat{p}) \rightsquigarrow \mathtt{inp}[\ell](p) : \mathtt{sum}[L, \ell](\{i \hookrightarrow \tau_i\}_{i \in L}; \ell \hookrightarrow \tau) \dashv\!\vert \hat{Y}} \quad (4.6e)$$

Again, the unexpanded and expanded pattern forms in the conclusion correspond and the premises correspond to those of the corresponding pattern typing rule, i.e. Rules (4.3a) through (4.3e), respectively. The upTSM context, $\hat{\Phi}$, passes through these rules opaquely. In Rule (4.6d), the conclusion of the rule collects all of the sigil expansions and hypotheses generated by the subpatterns. We define $\hat{Y}_i$ as shorthand for $\langle \mathcal{G}_i; Y_i \rangle$ and $\cup_{i \in L} \hat{Y}_i$ as shorthand for

$$\langle \cup_{i \in L} \mathcal{G}_i; \cup_{i \in L} Y_i \rangle$$

By the definition of iterated extension of finite functions, we implicitly have that no sigils or variables can be duplicated, i.e. that

$$\{\{\mathrm{dom}(\mathcal{G}_i) \cap \mathrm{dom}(\mathcal{G}_j) = \varnothing\}_{j \in L \setminus i}\}_{i \in L}$$

and

$$\{\{\mathrm{dom}(Y_i) \cap \mathrm{dom}(Y_j) = \varnothing\}_{j \in L \setminus i}\}_{i \in L}$$

The following lemma establishes that each well-typed expanded pattern can be expressed as an unexpanded pattern matching values of the same type and generating the same hypotheses and corresponding sigil updates. The metafunction $\mathcal{U}(Y)$ maps Y to an unexpanded typing context as follows:

$$\mathcal{U}(\varnothing) = \langle \varnothing; \varnothing \rangle$$
$$\mathcal{U}(Y, x : \tau) = \mathcal{U}(Y), \hat{x} \rightsquigarrow x : \tau$$
$$\mathcal{U}(\cup_{i \in L} Y_i) = \cup_{i \in L} \mathcal{U}(Y_i)$$

Lemma 4.12 (Pattern Expressibility). *If $\Delta \vdash p : \tau \dashv Y$ then $\Delta \vdash_{\hat{\Phi}} \mathcal{U}(p) \rightsquigarrow p : \tau \dashv \mathcal{U}(Y)$.*

Proof. By rule induction over Rules (4.3), using the definitions of $\mathcal{U}(Y)$ and $\mathcal{U}(p)$ given above. In each case, we can apply the IH to or over each premise, then apply the corresponding rule in Rules (4.6). $\qquad \square$

We can now establish the Expressibility Theorem – that each well-typed expanded expression, e, can be expressed as an unexpanded expression, $\hat{e}$, and assigned the same type under the corresponding contexts.

Theorem 4.13 (Expressibility). *Both of the following hold:*

1. *If $\Delta \, \Gamma \vdash e : \tau$ then $\mathcal{U}(\Delta) \, \mathcal{U}(\Gamma) \vdash_{\hat{\Psi}; \hat{\Phi}} \mathcal{U}(e) \rightsquigarrow e : \tau$.*
2. *If $\Delta \, \Gamma \vdash r : \tau \Rightarrow \tau'$ then $\mathcal{U}(\Delta) \, \mathcal{U}(\Gamma) \vdash_{\hat{\Psi}; \hat{\Phi}} \mathcal{U}(r) \rightsquigarrow r : \tau \Rightarrow \tau'$.*

Proof. By mutual rule induction over Rules (4.1) and Rule (4.2).

For part 1, we induct on the assumption. The rule transformation defined above guarantees that this part holds by its construction. In particular, in each case, we can apply Lemma 3.13 to or over each type formation premise, the IH (part 1) to or over each typing premise, the IH (part 2) over each rule typing premise, then apply the corresponding rule in Rules (4.4).

For part 2, we induct on the assumption. There is only one case:

Case (4.2). We have:

$$
\begin{array}{lll}
(1) & r = \texttt{rule}(p.e) & \text{by assumption} \\
(2) & \Delta \vdash p : \tau \dashv\!\!\vert\, Y & \text{by assumption} \\
(3) & \Delta\, \Gamma \cup Y \vdash e : \tau' & \text{by assumption} \\
(4) & \mathcal{U}(\Gamma) = \langle \mathcal{G}; \Gamma \rangle, \text{ for some } \mathcal{G} & \text{by definition of } \mathcal{U}(\Gamma) \\
(5) & \mathcal{U}(Y) = \langle \mathcal{G}'; Y \rangle, \text{ for some } \mathcal{G}' & \text{by definition of } \mathcal{U}(Y) \\
(6) & \mathcal{U}(\Gamma \cup Y) = \langle \mathcal{G} \uplus \mathcal{G}'; \Gamma \cup Y \rangle & \text{by definition of } \mathcal{U}(Y) \\
(7) & \mathcal{U}(r) = \texttt{urule}(\mathcal{U}(p).\mathcal{U}(e)) & \text{by definition of } \mathcal{U}(r) \\
(8) & \Delta \vdash_{\hat{\Phi}} \mathcal{U}(p) \rightsquigarrow p : \tau \dashv\!\!\vert\, \langle \mathcal{G}'; Y \rangle & \text{by Lemma 4.12 on (2)} \\
(9) & \hat{\Delta}\, \langle \mathcal{G} \uplus \mathcal{G}'; \Gamma \cup Y \rangle \vdash_{\hat{\Psi};\hat{\Phi}} \mathcal{U}(e) \rightsquigarrow e : \tau' & \text{by IH, part 1 on (3)} \\
(10) & \mathcal{U}(\Delta)\, \langle \mathcal{G}; \Gamma \rangle \vdash_{\hat{\Psi};\hat{\Phi}} \texttt{urule}(\mathcal{U}(p).\mathcal{U}(e)) \rightsquigarrow \texttt{rule}(p.e) : \tau \mapsto \tau' & \\
& & \text{by Rule (4.5) on (8)} \\
& & \text{and (9)}
\end{array}
$$

$\square$

ueTSM Definition and Application Rules (4.4l) and (4.4m) define typed expansion of ueTSM definitions and ueTSM application, respectively.

$$
\dfrac{\hat{\Delta} \vdash \hat{\tau} \rightsquigarrow \tau\ \text{type} \qquad \varnothing\, \varnothing \vdash e_{\text{parse}} : \texttt{parr}(\texttt{Body}; \texttt{ParseResultExp}) \qquad \hat{\Delta}\, \hat{\Gamma} \vdash_{\hat{\Psi},\hat{a}\rightsquigarrow a \hookrightarrow \texttt{uetsm}(\tau; e_{\text{parse}});\hat{\Phi}} \hat{e} \rightsquigarrow e : \tau'}{\hat{\Delta}\, \hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \texttt{syntax}\ \hat{a}\ \texttt{at}\ \hat{\tau}\ \texttt{for expressions}\ \{e_{\text{parse}}\}\ \texttt{in}\ \hat{e} \rightsquigarrow e : \tau'} \tag{4.4l}
$$

$$
\dfrac{b \downarrow e_{\text{body}} \qquad e_{\text{parse}}(e_{\text{body}}) \Downarrow \texttt{inj Success} \cdot e_{\text{cand}} \qquad e_{\text{cand}} \uparrow_{\texttt{CEExp}} \grave{e} \qquad \varnothing\, \varnothing \vdash^{\hat{\Delta};\hat{\Gamma};\hat{\Psi},\hat{a}\rightsquigarrow a \hookrightarrow \texttt{uetsm}(\tau; e_{\text{parse}});\hat{\Phi};b} \grave{e} \rightsquigarrow e : \tau}{\hat{\Delta}\, \hat{\Gamma} \vdash_{\hat{\Psi},\hat{a}\rightsquigarrow a \hookrightarrow \texttt{uetsm}(\tau; e_{\text{parse}});\hat{\Phi}} \hat{a}\ /b/ \rightsquigarrow e : \tau} \tag{4.4m}
$$

These rules are nearly identical to Rules (3.4l) and (3.4m), respectively, differing only in that the upTSM context, $\hat{\Phi}$, passes through them opaquely. The premises of these rules, and the following auxiliary definitions and conditions, can be understood as described in Sec. 3.2.6 and 3.2.7.

ueTSM contexts, $\hat{\Psi}$, are of the form $\langle \mathcal{A}; \Psi \rangle$, where $\mathcal{A}$ is a *TSM naming context* and Ψ is a *ueTSM definition context*. TSM naming contexts were defined in Sec. 3.2.6. A *ueTSM definition context*, Ψ, is a finite function mapping each symbol $a \in \text{dom}(\Psi)$ to an *expanded ueTSM definition*, $a \hookrightarrow \texttt{uetsm}(\tau; e_{\text{parse}})$, where τ is the ueTSM's type annotation, and e_{parse} is its parse function. We write $\Psi, a \hookrightarrow \texttt{uetsm}(\tau; e_{\text{parse}})$ when $a \notin \text{dom}(\Psi)$ for the extension of Ψ that maps a to $a \hookrightarrow \texttt{uetsm}(\tau; e_{\text{parse}})$. We write $\Delta \vdash \Psi$ ueTSMs when all the type annotations in Ψ are well-formed assuming Δ, and the parse functions in Ψ are closed and of type $\texttt{Body} \rightharpoonup \texttt{ParseResultExp}$.

Definition 4.14 (ueTSM Definition Context Formation). $\Delta \vdash \Psi$ ueTSMs *iff for each* $\hat{a} \hookrightarrow$ *uetsm*($\tau; e_{parse}$) $\in \Psi$, *we have* $\Delta \vdash \tau$ type *and* $\varnothing\, \varnothing \vdash e_{parse} : \texttt{Body} \rightharpoonup \texttt{ParseResultExp}$.

We define $\hat{\Psi}, \hat{a} \rightsquigarrow a \hookrightarrow \texttt{uetsm}(\tau; e_{\text{parse}})$, when $\hat{\Psi} = \langle \mathcal{A}; \Psi \rangle$, as an abbreviation of

$$
\langle \mathcal{A} \uplus \hat{a} \rightsquigarrow a; \Psi, a \hookrightarrow \texttt{uetsm}(\tau; e_{\text{parse}}) \rangle
$$

The type abbreviated `Body` classifies encodings of literal bodies, b. The mapping from literal bodies to values of type `Body` is defined by the *body encoding judgement* $b \downarrow e_{\text{body}}$. An inverse mapping is defined by the *body decoding judgement* $e_{\text{body}} \uparrow b$.

Judgement Form	Description
$b \downarrow e$	b has encoding e
$e \uparrow b$	e has decoding b

The following condition establishes an isomorphism between literal bodies and values of type `Body` mediated by the judgements above.

Condition 4.15 (Body Isomorphism). *All of the following must hold:*

1. *For every literal body b, we have that $b \downarrow e_{body}$ for some e_{body} such that $\vdash e_{body} :$ `Body` and e_{body} val.*
2. *If $\vdash e_{body} :$ `Body` and e_{body} val then $e_{body} \uparrow b$ for some b.*
3. *If $b \downarrow e_{body}$ then $e_{body} \uparrow b$.*
4. *If $\vdash e_{body} :$ `Body` and e_{body} val and $e_{body} \uparrow b$ then $b \downarrow e_{body}$.*
5. *If $b \downarrow e_{body}$ and $b \downarrow e'_{body}$ then $e_{body} = e'_{body}$.*
6. *If $\vdash e_{body} :$ `Body` and e_{body} val and $e_{body} \uparrow b$ and $e_{body} \uparrow b'$ then $b = b'$.*

`ParseResultExp` abbreviates a labeled sum type that distinguishes successful parses from parse errors:

$$\texttt{ParseResultExp} \triangleq [\texttt{Success} \hookrightarrow \texttt{CEExp}, \texttt{ParseError} \hookrightarrow \langle \rangle]$$

The type abbreviated `CEExp` classifies encodings of *candidate expansion expressions* (or *ce-expressions*), $\grave{e}$ (pronounced "grave e"). The syntax of ce-expressions will be described in Sec. 4.2.8. The mapping from ce-expressions to values of type `CEExp` is defined by the *ce-expression encoding judgement*, $\grave{e} \downarrow_{\text{CEExp}} e$. An inverse mapping is defined by the *ce-expression decoding judgement*, $e \uparrow_{\text{CEExp}} \grave{e}$.

Judgement Form	Description
$\grave{e} \downarrow_{\text{CEExp}} e$	$\grave{e}$ has encoding e
$e \uparrow_{\text{CEExp}} \grave{e}$	e has decoding $\grave{e}$

The following condition establishes an isomorphism between values of type `CEExp` and ce-expressions.

Condition 4.16 (Candidate Expansion Expression Isomorphism). *All of the following hold:*

1. *For every $\grave{e}$, we have $\grave{e} \downarrow_{\text{CEExp}} e_{cand}$ for some e_{cand} such that $\vdash e_{cand} :$ `CEExp` and e_{cand} val.*
2. *If $\vdash e_{cand} :$ `CEExp` and e_{cand} val then $e_{cand} \uparrow_{\text{CEExp}} \grave{e}$ for some $\grave{e}$.*
3. *If $\grave{e} \downarrow_{\text{CEExp}} e_{cand}$ then $e_{cand} \uparrow_{\text{CEExp}} \grave{e}$.*
4. *If $\vdash e_{cand} :$ `CEExp` and e_{cand} val and $e_{cand} \uparrow_{\text{CEExp}} \grave{e}$ then $\grave{e} \downarrow_{\text{CEExp}} e_{cand}$.*
5. *If $\grave{e} \downarrow_{\text{CEExp}} e_{cand}$ and $\grave{e} \downarrow_{\text{CEExp}} e'_{cand}$ then $e_{cand} = e'_{cand}$.*
6. *If $\vdash e_{cand} :$ `CEExp` and e_{cand} val and $e_{cand} \uparrow_{\text{CEExp}} \grave{e}$ and $e_{cand} \uparrow_{\text{CEExp}} \grave{e}'$ then $\grave{e} = \grave{e}'$.*

upTSM Definition and Application   Rules (4.4n) and (4.6f) define upTSM definition and application, and are defined in the next two subsections, respectively.

4.2.6 upTSM Definition

The stylized upTSM definition form is

$$\texttt{syntax } \hat{a} \texttt{ at } \hat{\tau} \texttt{ for patterns } \{e_{\mathrm{parse}}\} \texttt{ in } \hat{e}$$

An unexpanded expression of this form defines a upTSM named $\hat{a}$ with *unexpanded type annotation* $\hat{\tau}$ and *parse function* e_{parse} for use within $\hat{e}$.

Rule (4.4n) defines typed expansion of upTSM definitions:

$$\frac{\begin{array}{cc} \hat{\Delta} \vdash \hat{\tau} \rightsquigarrow \tau \text{ type} & \varnothing \, \varnothing \vdash e_{\mathrm{parse}} : \texttt{Body} \rightharpoonup \texttt{ParseResultPat} \\ \hat{\Delta} \, \hat{\Gamma} \vdash_{\hat{\Psi}; \hat{\Phi}, \hat{a} \rightsquigarrow a \hookrightarrow \texttt{uptsm}(\tau; e_{\mathrm{parse}})} \hat{e} \rightsquigarrow e : \tau' \end{array}}{\hat{\Delta} \, \hat{\Gamma} \vdash_{\hat{\Psi}; \hat{\Phi}} \texttt{syntax } \hat{a} \texttt{ at } \hat{\tau} \texttt{ for patterns } \{e_{\mathrm{parse}}\} \texttt{ in } \hat{e} \rightsquigarrow e : \tau'} \quad (4.4\text{n})$$

This rule is similar to Rule (4.4l), which governs ueTSM definitions. The premises of this rule can be understood as follows, in order:

1. The first premise ensures that the unexpanded type annotation is well-formed and expands it to produce the *type annotation*, τ.

2. The second premise checks that the parse function, e_{parse}, is closed and of type

$$\texttt{Body} \rightharpoonup \texttt{ParseResultExp}$$

The type abbreviated `Body` is characterized above.

`ParseResultPat`, like `ParseResultExp` above, abbreviates a labeled sum type that distinguishes successful parses from parse errors:

$$\texttt{ParseResultPat} \triangleq [\texttt{Success} \hookrightarrow \texttt{CEPat}, \texttt{ParseError} \hookrightarrow \langle\rangle]$$

The type abbreviated `CEPat` classifies encodings of *candidate expansion patterns* (or *ce-patterns*), $\grave{p}$. The syntax of ce-patterns will be described in Sec. 4.2.8. The mapping from ce-patterns to values of type `CEPat` is defined by the *ce-pattern encoding judgement*, $\grave{p} \downarrow_{\mathrm{CEPat}} e$. An inverse mapping is defined by the *ce-pattern decoding judgement*, $e \uparrow_{\mathrm{CEPat}} \grave{p}$.

Judgement Form	Description
$\grave{p} \downarrow_{\mathrm{CEPat}} e$	$\grave{p}$ has encoding e
$e \uparrow_{\mathrm{CEPat}} \grave{p}$	e has decoding $\grave{p}$

Again, rather than picking a particular definition of `CEPat` and defining the judgements above inductively against it, we only state the following condition, which establishes an isomorphism between values of type `CEPat` and ce-patterns.

Condition 4.17 (Candidate Expansion Pattern Isomorphism)**.** *All of the following must hold:*

   (a) *For every $\grave{p}$, we have $\grave{p} \downarrow_{\mathrm{CEPat}} e_{cand}$ for some e_{cand} such that $\vdash e_{cand} : \texttt{CEPat}$ and e_{cand} val.*

(b) If $\vdash e_{cand}$: CEPat and e_{cand} val then $e_{cand} \uparrow_{\mathsf{CEPat}} \grave{p}$ for some $\grave{p}$.

(c) If $\grave{p} \downarrow_{\mathsf{CEPat}} e_{cand}$ then $e_{cand} \uparrow_{\mathsf{CEPat}} \grave{p}$.

(d) If $\vdash e_{cand}$: CEPat and e_{cand} val and $e_{cand} \uparrow_{\mathsf{CEPat}} \grave{p}$ then $\grave{p} \downarrow_{\mathsf{CEPat}} e_{cand}$.

(e) If $\grave{p} \downarrow_{\mathsf{CEPat}} e_{cand}$ and $\grave{p} \downarrow_{\mathsf{CEPat}} e'_{cand}$ then $e_{cand} = e'_{cand}$.

(f) If $\vdash e_{cand}$: CEPat and e_{cand} val and $e_{cand} \uparrow_{\mathsf{CEPat}} \grave{p}$ and $e_{cand} \uparrow_{\mathsf{CEPat}} \grave{p}'$ then $\grave{p} = \grave{p}'$.

3. The final premise of Rule (4.4n) extends the upTSM context, $\hat{\Phi}$, with the newly determined upTSM definition, and proceeds to assign a type, τ', and expansion, e, to $\hat{e}$. The conclusion of Rule (4.4n) assigns this type and expansion to the upTSM definition as a whole.

upTSM contexts, $\hat{\Phi}$, are of the form $\langle \mathcal{A}; \Phi \rangle$, where $\mathcal{A}$ is a TSM naming context, defined previously, and Φ is a *upTSM definition context*.

A *upTSM definition context*, Φ, is a finite function mapping each symbol $a \in \mathrm{dom}(\Phi)$ to an *expanded upTSM definition*, $a \hookrightarrow \mathtt{uptsm}(\tau; e_{\mathrm{parse}})$, where τ is the upTSM's type annotation, and e_{parse} is its parse function. We write $\Phi, a \hookrightarrow \mathtt{uptsm}(\tau; e_{\mathrm{parse}})$ when $a \notin \mathrm{dom}(\Phi)$ for the extension of Φ that maps a to $a \hookrightarrow \mathtt{uptsm}(\tau; e_{\mathrm{parse}})$. We write $\Delta \vdash \Phi$ upTSMs when all the type annotations in Φ are well-formed assuming Δ, and the parse functions in Φ are closed and of type $\mathtt{Body} \rightharpoonup \mathtt{ParseResultPat}$.

Definition 4.18 (upTSM Definition Context Formation). $\Delta \vdash \Phi$ upTSMs *iff for each $a \hookrightarrow \mathtt{uptsm}(\tau; e_{parse}) \in \Phi$, we have $\Delta \vdash \tau$ type and $\varnothing\,\varnothing \vdash e_{parse}$: $\mathtt{Body} \rightharpoonup \mathtt{ParseResultPat}$.*

We define $\hat{\Phi}, \hat{a} \rightsquigarrow a \hookrightarrow \mathtt{uptsm}(\tau; e_{\mathrm{parse}})$, when $\hat{\Phi} = \langle \mathcal{A}; \Phi \rangle$, as an abbreviation of

$$\langle \mathcal{A} \uplus \hat{a} \rightsquigarrow a; \Phi, a \hookrightarrow \mathtt{uptsm}(\tau; e_{\mathrm{parse}}) \rangle$$

and $\hat{\Phi} \cup \hat{\Phi}'$ when $\hat{\Phi} = \langle \mathcal{A}; \Phi \rangle$ and $\hat{\Phi}' = \langle \mathcal{A}'; \Phi' \rangle$ as an abbreviation of

$$\langle \mathcal{A} \cup \mathcal{A}'; \Phi \cup \Phi' \rangle$$

4.2.7 upTSM Application

The stylized unexpanded pattern form for applying a upTSM named $\hat{a}$ to a literal form with literal body b is:

$$\hat{a} \ /b/$$

This stylized form is identical to the stylized form for ueTSM application, differing in that appears within the syntax of unexpanded patterns, $\hat{p}$, rather than unexpanded expressions, $\hat{e}$. The corresponding operational form is $\mathtt{uapuptsm}[b][\hat{a}]$.

Rule (4.6f), below, governs upTSM application.

$$\frac{\begin{array}{ccc} b \downarrow e_{\mathrm{body}} & e_{\mathrm{parse}}(e_{\mathrm{body}}) \Downarrow \mathtt{inj\ Success} \cdot e_{\mathrm{cand}} & e_{\mathrm{cand}} \uparrow_{\mathsf{CEPat}} \grave{p} \\ \multicolumn{3}{c}{\vdash^{\Delta; \hat{\Phi}, \hat{a} \rightsquigarrow a \hookrightarrow \mathtt{uptsm}(\tau; e_{\mathrm{parse}}); b} \grave{p} \rightsquigarrow p : \tau \dashv\!|^{\hat{Y}}} \end{array}}{\Delta \vdash_{\hat{\Phi}, \hat{a} \rightsquigarrow a \hookrightarrow \mathtt{uptsm}(\tau; e_{\mathrm{parse}})} \hat{a} \ /b/ \rightsquigarrow p : \tau \dashv\!| \hat{Y}} \tag{4.6f}$$

This rule is similar to Rule (4.4m), which governs ueTSM application. Its premises can be understood as follows, in order:

1. The first premise determines the encoding of the literal body, e_{body}.

2. The second premise applies the parse function e_{parse} to e_{body}. If parsing succeeds, i.e. a value of the (stylized) form $\mathtt{inj}\ \mathtt{Success} \cdot e_{\text{cand}}$ results from evaluation, then e_{cand} will be a value of type $\mathtt{CEPat}$ (assuming a well-formed upTSM context, by application of Assumption 4.9). We call e_{cand} the *encoding of the candidate expansion*.

3. The third premise decodes the encoding of the candidate expansion to produce *candidate expansion*, $\grave{p}$.

4. The final premise *validates* the candidate expansion and simultaneously generates the final expansion, p, and assumptions, Y. This is the topic of Sec. 4.2.9.

4.2.8 Syntax of Candidate Expansions

Figure 4.4 defines the syntax of candidate expansion types (or *ce-types*), $\grave{\tau}$, candidate expansion expressions (or *ce-expressions*), $\grave{e}$, candidate expansion rules (or *ce-rules*), $\grave{r}$, and candidate expansion patterns (or *ce-patterns*), $\grave{p}$. Candidate expansion terms are identified up to α-equivalence in the usual manner.

Each inner core form, except for the variable pattern form, maps onto a candidate expansion form. We refer to these as the *shared forms*. In particular:

- Each type form maps onto a ce-type form according to the metafunction $\mathcal{C}(\tau)$, defined in Sec. 3.2.8.

- Each expanded expression form maps onto a ce-expression form according to the metafunction $\mathcal{C}(e)$, defined as follows:

$$\mathcal{C}(x) = x$$
$$\mathcal{C}(\mathtt{lam}\{\tau\}(x.e)) = \mathtt{celam}\{\mathcal{C}(\tau)\}(x.\mathcal{C}(e))$$
$$\mathcal{C}(\mathtt{ap}(e_1;e_2)) = \mathtt{ceap}(\mathcal{C}(e_1);\mathcal{C}(e_2))$$
$$\mathcal{C}(\mathtt{tlam}(t.e)) = \mathtt{cetlam}(t.\mathcal{C}(e))$$
$$\mathcal{C}(\mathtt{tap}\{\tau\}(e)) = \mathtt{cetap}\{\mathcal{C}(\tau)\}(\mathcal{C}(e))$$
$$\mathcal{C}(\mathtt{fold}\{t.\tau\}(e)) = \mathtt{cefold}\{t.\mathcal{C}(\tau)\}(\mathcal{C}(e))$$
$$\mathcal{C}(\mathtt{unfold}(e)) = \mathtt{ceunfold}(\mathcal{C}(e))$$
$$\mathcal{C}(\mathtt{tpl}[L](\{i \hookrightarrow e_i\}_{i\in L})) = \mathtt{cetpl}[L](\{i \hookrightarrow \mathcal{C}(e_i)\}_{i\in L})$$
$$\mathcal{C}(\mathtt{in}[L][\ell]\{\{i \hookrightarrow \tau_i\}_{i\in L}\}(e)) = \mathtt{cein}[L][\ell]\{\{i \hookrightarrow \mathcal{C}(\tau_i)\}_{i\in L}\}(\mathcal{C}(e))$$
$$\mathcal{C}(\mathtt{match}[n]\{\tau\}(e;\{r_i\}_{1\leq i\leq n})) = \mathtt{cematch}[n]\{\mathcal{C}(\tau)\}(\mathcal{C}(e);\{\mathcal{C}(r_i)\}_{1\leq i\leq n})$$

- The expanded rule form maps onto the ce-rule form according to the metafunction $\mathcal{C}(r)$, defined as follows:

$$\mathcal{C}(\mathtt{rule}(p.e)) = \mathtt{cerule}(p.\mathcal{C}(e))$$

Sort			Operational Form	Stylized Form	Description
CETyp	$\check{\tau}$	$::=$	t	t	variable
			$\mathtt{ceparr}(\check{\tau};\check{\tau})$	$\check{\tau} \rightharpoonup \check{\tau}$	partial function
			$\mathtt{ceall}(t.\check{\tau})$	$\forall t.\check{\tau}$	polymorphic
			$\mathtt{cerec}(t.\check{\tau})$	$\mu t.\check{\tau}$	recursive
			$\mathtt{ceprod}[L](\{i \hookrightarrow \check{\tau}_i\}_{i \in L})$	$\langle\{i \hookrightarrow \check{\tau}_i\}_{i \in L}\rangle$	labeled product
			$\mathtt{cesum}[L](\{i \hookrightarrow \check{\tau}_i\}_{i \in L})$	$[\{i \hookrightarrow \check{\tau}_i\}_{i \in L}]$	labeled sum
			$\mathtt{cesplicedt}[m;n]$	$\mathtt{splicedt}\langle m,n\rangle$	spliced
CEExp	$\check{e}$	$::=$	x	x	variable
			$\mathtt{celam}\{\check{\tau}\}(x.\check{e})$	$\lambda x{:}\check{\tau}.\check{e}$	abstraction
			$\mathtt{ceap}(\check{e};\check{e})$	$\check{e}(\check{e})$	application
			$\mathtt{cetlam}(t.\check{e})$	$\Lambda t.\check{e}$	type abstraction
			$\mathtt{cetap}\{\check{\tau}\}(\check{e})$	$\check{e}[\check{\tau}]$	type application
			$\mathtt{cefold}\{t.\check{\tau}\}(\check{e})$	$\mathtt{fold}(\check{e})$	fold
			$\mathtt{ceunfold}(\check{e})$	$\mathtt{unfold}(\check{e})$	unfold
			$\mathtt{cetpl}[L](\{i \hookrightarrow \check{e}_i\}_{i \in L})$	$\langle\{i \hookrightarrow \check{e}_i\}_{i \in L}\rangle$	labeled tuple
			$\mathtt{cepr}[\ell](\check{e})$	$\check{e} \cdot \ell$	projection
			$\mathtt{cein}[L][\ell]\{\{i \hookrightarrow \check{\tau}_i\}_{i \in L}\}(\check{e})$	$\mathtt{inj}\ \ell \cdot \check{e}$	injection
			$\mathtt{cematch}[n]\{\check{\tau}\}(\check{e};\{\check{r}_i\}_{1 \leq i \leq n})$	$\mathtt{match}\ \check{e}\ \{\check{r}_i\}_{1 \leq i \leq n}$	match
			$\mathtt{cesplicede}[m;n]$	$\mathtt{splicede}\langle m,n\rangle$	spliced
CERule	$\check{r}$	$::=$	$\mathtt{cerule}(p.\check{e})$	$p \Rightarrow \check{e}$	rule
CEPat	$\check{p}$	$::=$	$\mathtt{cewildp}$	$\_$	wildcard pattern
			$\mathtt{cefoldp}(p)$	$\mathtt{fold}(p)$	fold pattern
			$\mathtt{cetplp}[L](\{i \hookrightarrow \check{p}_i\}_{i \in L})$	$\langle\{i \hookrightarrow \check{p}_i\}_{i \in L}\rangle$	labeled tuple pattern
			$\mathtt{ceinp}[\ell](\check{p})$	$\mathtt{inj}\ \ell \cdot \check{p}$	injection pattern
			$\mathtt{cesplicedp}[m;n]$	$\mathtt{splicedp}\langle m,n\rangle$	spliced

Figure 4.4: Abstract syntax of candidate expansion types, expressions, rules and patterns in miniVerse$_U$. Candidate expansion terms are identified up to α-equivalence.

Notice that ce-rules bind expanded patterns, not ce-patterns. This is because ce-rules appear in ce-expressions, which are generated by ueTSMs. It would not be sensible for a ueTSM to splice a pattern out of a literal body.

- Each expanded pattern form, except for the variable pattern form, maps onto a ce-pattern form according to the metafunction $\mathcal{C}(p)$, defined as follows:

$$\mathcal{C}(\texttt{wildp}) = \texttt{cewildp}$$
$$\mathcal{C}(\texttt{foldp}(p)) = \texttt{cefoldp}(\mathcal{C}(p))$$
$$\mathcal{C}(\texttt{tplp}[L](\{i \hookrightarrow p_i\}_{i \in L})) = \texttt{cetplp}[L](\{i \hookrightarrow \mathcal{C}(p_i)\}_{i \in L})$$
$$\mathcal{C}(\texttt{inp}[\ell](p)) = \texttt{ceinp}[\ell](\mathcal{C}(p))$$

There are three other candidate expansion forms: a ce-type form for *references to spliced unexpanded types*, $\texttt{cesplicedt}[m; n]$, a ce-expression form for *references to spliced unexpanded expressions*, $\texttt{cesplicede}[m; n]$, and, highlighted in gray in Figure 4.4, a ce-pattern form for *references to spliced unexpanded patterns*, $\texttt{cesplicedp}[m; n]$.

4.2.9 Candidate Expansion Validation

The *candidate expansion validation judgements* validate ce-terms and simultaneously generate their final expansions.

Judgement Form	Description
$\Delta \vdash^{\mathbb{T}} \grave{\tau} \rightsquigarrow \tau$ type	$\grave{\tau}$ is well-formed and has expansion τ assuming Δ and type splicing scene $\mathbb{T}$
$\Delta \, \Gamma \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e : \tau$	$\grave{e}$ has expansion e and type τ assuming Δ and Γ and expression splicing scene $\mathbb{E}$
$\Delta \, \Gamma \vdash^{\mathbb{E}} \grave{r} \rightsquigarrow r : \tau \mapsto \tau'$	$\grave{r}$ has expansion r and takes values of type τ to values of type τ' assuming Δ and Γ and expression splicing scene $\mathbb{E}$
$\vdash^{\mathbb{P}} \grave{p} \rightsquigarrow p : \tau \dashv\mid^{\hat{Y}}$	$\grave{p}$ expands to p and matches values of type τ generating assumptions $\hat{Y}$ assuming pattern splicing scene $\mathbb{P}$

Expression splicing scenes, $\mathbb{E}$, are of the form $\hat{\Delta}; \hat{\Gamma}; \hat{\Psi}; \hat{\Phi}; b$, *type splicing scenes*, $\mathbb{T}$, are of the form $\hat{\Delta}; b$, and *pattern splicing scenes*, $\mathbb{P}$, are of the form $\hat{\Delta}; \hat{\Phi}; b$. Their purpose is to "remember", during candidate expansion validation, the contexts, TSM environments and literal bodies from the TSM application site (cf. Rules (4.4m) and (4.6f)), because these are necessary to validate references to spliced terms. We write $\texttt{ts}(\mathbb{E})$ for the type splicing scene constructed by dropping the unexpanded typing context and TSM environments from $\mathbb{E}$:

$$\texttt{ts}(\hat{\Delta}; \hat{\Gamma}; \hat{\Psi}; \hat{\Phi}; b) = \hat{\Delta}; b$$

Candidate Expansion Type Validation

The *ce-type validation judgement*, $\Delta \vdash^{\mathbb{T}} \grave{\tau} \rightsquigarrow \tau$ type, is inductively defined by Rules (3.5), which were defined in Sec. 3.2.9.

Candidate Expansion Expression and Rule Validation

The *ce-expression validation judgement*, $\Delta\,\Gamma \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e : \tau$, and the *ce-rule validation judgement*, $\Delta\,\Gamma \vdash^{\mathbb{E}} \grave{r} \rightsquigarrow r : \tau \mapsto \tau'$, are defined mutually inductively with Rules (4.4) and Rule (4.5) by Rules (4.7) and Rule (4.8), respectively, as follows.

Rules (4.7) define ce-expression validation and consist of the following rules:

- Rules written identically to Rules (3.6a) through (3.6j). We will refer to these as Rules (4.7a) through (4.7j).

- The following rule for match ce-expressions:

$$\frac{\Delta\,\Gamma \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e : \tau \qquad \Delta \vdash^{\mathsf{ts}(\mathbb{E})} \grave{\tau}' \rightsquigarrow \tau' \text{ type} \qquad \{\Delta\,\Gamma \vdash^{\mathbb{E}} \grave{r}_i \rightsquigarrow r_i : \tau \mapsto \tau'\}_{1 \le i \le n}}{\Delta\,\Gamma \vdash^{\mathbb{E}} \mathtt{cematch}[n]\{\grave{\tau}'\}(\grave{e}; \{\grave{r}_i\}_{1 \le i \le n}) \rightsquigarrow \mathtt{match}[n]\{\tau'\}(e; \{r_i\}_{1 \le i \le n}) : \tau'} \quad (4.7\text{k})$$

- The following rule for references to spliced unexpanded expressions, which can be understood as described in Sec. 3.2.9.

$$\frac{\mathsf{parseUExp}(\mathsf{subseq}(b; m; n)) = \hat{e} \qquad \langle \mathcal{D}; \Delta_{\mathsf{app}} \rangle\, \langle \mathcal{G}; \Gamma_{\mathsf{app}} \rangle \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{e} \rightsquigarrow e : \tau \qquad \Delta \cap \Delta_{\mathsf{app}} = \varnothing \qquad \mathsf{dom}(\Gamma) \cap \mathsf{dom}(\Gamma_{\mathsf{app}}) = \varnothing}{\Delta\,\Gamma \vdash^{\langle \mathcal{D}; \Delta_{\mathsf{app}} \rangle; \langle \mathcal{G}; \Gamma_{\mathsf{app}} \rangle; \hat{\Psi}; \hat{\Phi}; b} \mathtt{cesplicede}[m; n] \rightsquigarrow e : \tau} \quad (4.7\text{l})$$

Rule (4.8) defines ce-rule validation and is defined as follows:

$$\frac{\Delta \vdash p : \tau \dashv\! Y \qquad \Delta\,\Gamma \cup Y \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e : \tau'}{\Delta\,\Gamma \vdash^{\mathbb{E}} \mathtt{cerule}(p.\grave{e}) \rightsquigarrow \mathtt{rule}(p.e) : \tau \mapsto \tau'} \quad (4.8)$$

The following lemma establishes that each well-typed expanded expression, e, can be expressed as a valid ce-expression, $\mathcal{C}(e)$, that is assigned the same type under any expression splicing scene.

Theorem 4.19 (Candidate Expansion Expression Expressibility)**.** *Both of the following hold:*

1. *If $\Delta\,\Gamma \vdash e : \tau$ then $\Delta\,\Gamma \vdash^{\mathbb{E}} \mathcal{C}(e) \rightsquigarrow e : \tau$.*
2. *If $\Delta\,\Gamma \vdash r : \tau \mapsto \tau'$ then $\Delta\,\Gamma \vdash^{\mathbb{E}} \mathcal{C}(r) \rightsquigarrow r : \tau \mapsto \tau'$.*

Proof. By mutual rule induction over Rules (4.1) and Rule (4.2).

For part 1, we induct on the assumption.

Case (4.1a) through (4.1j)**.** In each of these cases, we apply Lemma 3.18 to or over each type formation premise, the IH (part 1) to or over each typing premise, then apply the corresponding ce-expression validation rule in Rules (4.7a) through (4.7j).

Case (4.1k)**.** We have:

$$
\begin{array}{lll}
(1) & e = \mathtt{match}[n]\{\tau\}(e'; \{r_i\}_{1 \le i \le n}) & \text{by assumption} \\
(2) & \mathcal{C}(e) = \mathtt{cematch}[n]\{\mathcal{C}(\tau)\}(\mathcal{C}(e'); \{\mathcal{C}(r_i)\}_{1 \le i \le n}) & \text{by definition of } \mathcal{C}(e) \\
(3) & \Delta\,\Gamma \vdash e' : \tau' & \text{by assumption}
\end{array}
$$

(4) $\Delta \vdash \tau$ type	by assumption
(5) $\{\Delta \, \Gamma \vdash r_i : \tau' \Mapsto \tau\}_{1 \le i \le n}$	by assumption
(6) $\Delta \, \Gamma \vdash^{\mathbb{E}} \mathcal{C}(e') \rightsquigarrow e' : \tau'$	by IH, part 1 on (3)
(7) $\Delta \vdash^{\mathsf{ts}(\mathbb{E})} \mathcal{C}(\tau) \rightsquigarrow \tau$ type	by Lemma 3.19 on (4)
(8) $\{\Delta \, \Gamma \vdash^{\mathbb{E}} \mathcal{C}(r_i) \rightsquigarrow r_i : \tau' \Mapsto \tau\}_{1 \le i \le n}$	by IH, part 2 over (5)
(9) $\Delta \, \Gamma \vdash^{\mathbb{E}} \mathtt{cematch}[n]\{\mathcal{C}(\tau)\}(\mathcal{C}(e'); \{\mathcal{C}(r_i)\}_{1 \le i \le n}) \rightsquigarrow$	
$\quad \mathtt{match}[n]\{\tau\}(e'; \{r_i\}_{1 \le i \le n}) : \tau$	by Rule (4.7k) on (6), (7) and (8)

For part 2, we induct on the assumption. There is only one case.

Case (4.2). We have:

(1) $r = \mathtt{rule}(p.e)$	by assumption
(2) $\mathcal{C}(r) = \mathtt{cerule}(p.\mathcal{C}(e))$	by definition of $\mathcal{C}(r)$
(3) $\Delta \vdash p : \tau \dashv Y$	by assumption
(4) $\Delta \, \Gamma \cup Y \vdash e : \tau'$	by assumption
(5) $\Delta \, \Gamma \cup Y \vdash^{\mathbb{E}} \mathcal{C}(e) \rightsquigarrow e : \tau'$	by IH, part 1 on (4)
(6) $\Delta \, \Gamma \vdash^{\mathbb{E}} \mathtt{cerule}(p.\mathcal{C}(e)) \rightsquigarrow \mathtt{rule}(p.e) : \tau \Mapsto \tau'$	by Rule (4.8) on (3) and (5)

$\square$

Candidate Expansion Pattern Validation

upTSMs generate candidate expansions of ce-pattern form, as described in Sec. 4.2.7. The *ce-pattern validation judgement*, $\vdash^{\mathbb{P}} \grave{p} \rightsquigarrow p : \tau \dashv^{\hat{Y}}$, which appears as the final premise of Rule (4.4m), validates ce-patterns by checking that the pattern matches values of type τ, and simultaneously generates the final expansion, p, and the hypotheses $\hat{Y}$. Hypotheses can be generated only by spliced subpatterns, so there is no ce-pattern form corresponding to variable patterns (this is also why $\hat{Y}$ appears as a superscript). The pattern splicing scene, $\mathbb{P}$, is used to "remember" the upTSM context and literal body from the upTSM application site (cf. Rule (4.6f)).

The ce-pattern validation judgement is defined mutually inductively with Rules (4.6) by the following rules.

$$\frac{}{\vdash^{\mathbb{P}} \mathtt{cewildp} \rightsquigarrow \mathtt{wildp} : \tau \dashv^{\langle \varnothing; \varnothing \rangle}} \tag{4.9a}$$

$$\frac{\vdash^{\mathbb{P}} \grave{p} \rightsquigarrow p : [\mathtt{rec}(t.\tau)/t]\tau \dashv^{\hat{Y}}}{\vdash^{\mathbb{P}} \mathtt{cefoldp}(\grave{p}) \rightsquigarrow \mathtt{foldp}(p) : \mathtt{rec}(t.\tau) \dashv^{\hat{Y}}} \tag{4.9b}$$

$$\frac{\{\vdash^{\mathbb{P}} \grave{p}_i \rightsquigarrow p_i : \tau_i \dashv\!\!\mid^{\hat{Y}_i}\}_{i \in L}}{\left(\begin{array}{c} \vdash^{\mathbb{P}} \texttt{cetplp}[L](\{i \hookrightarrow \grave{p}_i\}_{i \in L}) \\ \rightsquigarrow \\ \texttt{tplp}[L](\{i \hookrightarrow p_i\}_{i \in L}) : \texttt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L}) \dashv\!\!\mid^{\cup_{i \in L} \hat{Y}_i} \end{array}\right)} \quad (4.9c)$$

$$\frac{\vdash^{\mathbb{P}} \grave{p} \rightsquigarrow p : \tau \dashv\!\!\mid^{\hat{Y}}}{\vdash^{\mathbb{P}} \texttt{ceinp}[\ell](\grave{p}) \rightsquigarrow \texttt{inp}[\ell](p) : \texttt{sum}[L, \ell](\{i \hookrightarrow \tau_i\}_{i \in L}; \ell \hookrightarrow \tau) \dashv\!\!\mid^{\hat{Y}}} \quad (4.9d)$$

$$\frac{\mathsf{parseUPat}(\mathsf{subseq}(b; m; n)) = \hat{p} \qquad \Delta \vdash_{\hat{\Phi}} \hat{p} \rightsquigarrow p : \tau \dashv\!\!\mid \hat{Y}}{\vdash^{\Delta; \hat{\Phi}; b} \texttt{cesplicedp}[m; n] \rightsquigarrow p : \tau \dashv\!\!\mid^{\hat{Y}}} \quad (4.9e)$$

Rules (4.9a) through (4.9d) handle ce-patterns of shared form, and correspond to Rules (4.3b) through (4.3e). Rule (4.9e) handles references to spliced unexpanded patterns. The first premise parses the indicated subsequence of the literal body, b, to produce the referenced unexpanded pattern, $\hat{p}$, and the second premise types and expands $\hat{p}$ under the upTSM context Φ from the upTSM application site, producing the hypotheses Y. These are the hypotheses generated in the conclusion of the rule.

Notice that none of these rules explicitly add any hypotheses to the pattern typing context, so upTSMs cannot introduce any hypotheses other than those that come from such spliced subpatterns. This achieves the "no hidden assumptions" hygiene property described in Sec. 4.1.5.

The following lemma establishes that every well-typed expanded pattern that generates no hypotheses can be expressed as a ce-pattern.

Lemma 4.20 (Candidate Expansion Pattern Expressibility). *If* $\Delta \vdash p : \tau \dashv\!\!\mid \varnothing$ *then* $\vdash^{\Delta; \hat{\Phi}; b} \mathcal{C}(p) \rightsquigarrow p : \tau \dashv\!\!\mid^{\langle \varnothing; \varnothing \rangle}$.

Proof. By rule induction over Rules (4.3).

Case (4.3a). This case does not apply.

Case (4.3b). We have:

(1)	$p = \texttt{wildp}$	by assumption
(2)	$\mathcal{C}(p) = \texttt{cewildp}$	by definition of $\mathcal{C}(p)$
(3)	$\vdash^{\Delta; \hat{\Phi}; b} \texttt{cewildp} \rightsquigarrow \texttt{wildp} : \tau \dashv\!\!\mid^{\langle \varnothing; \varnothing \rangle}$	by Rule (4.9a)

Case (4.3c). We have:

(1)	$p = \texttt{foldp}(p')$	by assumption
(2)	$\mathcal{C}(p) = \texttt{cefoldp}(\mathcal{C}(p'))$	by definition of $\mathcal{C}(p)$
(3)	$\tau = \texttt{rec}(t.\tau')$	by assumption
(4)	$\Delta \vdash p' : [\texttt{rec}(t.\tau')/t]\tau' \dashv\!\!\mid \varnothing$	by assumption
(5)	$\vdash^{\Delta; \hat{\Phi}; b} \mathcal{C}(p') \rightsquigarrow p : [\texttt{rec}(t.\tau')/t]\tau' \dashv\!\!\mid^{\langle \varnothing; \varnothing \rangle}$	by IH on (4)
(6)	$\vdash^{\Delta; \hat{\Phi}; b} \texttt{cefoldp}(\mathcal{C}(p')) \rightsquigarrow \texttt{foldp}(p') : \texttt{rec}(t.\tau') \dashv\!\!\mid^{\langle \varnothing; \varnothing \rangle}$	
		by Rule (4.9b) on (5)

Case (4.3d). We have:

$$
\begin{array}{lll}
(1) & p = \mathtt{tplp}[L](\{i \hookrightarrow p_i\}_{i \in L}) & \text{by assumption} \\
(2) & \mathcal{C}(p) = \mathtt{cetpl}[L](\{i \hookrightarrow \mathcal{C}(p_i)\}_{i \in L}) & \text{by definition of } \mathcal{C}(p) \\
(3) & \tau = \mathtt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L}) & \text{by assumption} \\
(4) & \{\Delta \vdash p_i : \tau_i \dashv \varnothing\}_{i \in L} & \text{by assumption} \\
(5) & \{\vdash^{\Delta;\hat{\Phi};b} \mathcal{C}(p_i) \rightsquigarrow p_i : \tau_i \dashv^{\langle \varnothing;\varnothing \rangle}\}_{i \in L} & \text{by IH over (4)} \\
(6) & \vdash^{\Delta;\hat{\Phi};b} \mathtt{cetpl}[L](\{i \hookrightarrow \mathcal{C}(p_i)\}_{i \in L}) \rightsquigarrow \mathtt{tplp}[L](\{i \hookrightarrow p_i\}_{i \in L}) : & \\
& \quad \mathtt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L}) \dashv^{\langle \varnothing;\varnothing \rangle} & \text{by Rule (4.9c) on (5)}
\end{array}
$$

Case (4.3e). We have:

$$
\begin{array}{lll}
(1) & p = \mathtt{inp}[\ell](p') & \text{by assumption} \\
(2) & \mathcal{C}(p) = \mathtt{ceinp}[\ell](\mathcal{C}(p')) & \text{by definition of } \mathcal{C}(p) \\
(3) & \tau = \mathtt{sum}[L, \ell](\{i \hookrightarrow \tau_i\}_{i \in L}; \ell \hookrightarrow \tau') & \text{by assumption} \\
(4) & \Delta \vdash p' : \tau' \dashv \varnothing & \text{by assumption} \\
(5) & \vdash^{\Delta;\hat{\Phi};b} \mathcal{C}(p') \rightsquigarrow p' : \tau' \dashv^{\langle \varnothing;\varnothing \rangle} & \text{by IH on (4)} \\
(6) & \vdash^{\Delta;\hat{\Phi};b} \mathtt{ceinp}[\ell](\mathcal{C}(p')) \rightsquigarrow \mathtt{inp}[\ell](p') : \mathtt{sum}[L, \ell](\{i \hookrightarrow \tau_i\}_{i \in L}; \ell \hookrightarrow & \\
& \quad \tau') \dashv^{\langle \varnothing;\varnothing \rangle} & \text{by Rule (4.9d) on (5)}
\end{array}
$$

$\square$

4.2.10 Metatheory

The following theorem establishes that typed pattern expansion produces an expanded pattern that matches values of the specified type and generates the same hypotheses. It must be stated mutually with the corresponding theorem about candidate expansion patterns, because the judgements are mutually defined.

Theorem 4.21 (Typed Pattern Expansion)**.** *Both of the following hold:*

1. *If* $\Delta \vdash_{\langle \mathcal{A}; \Phi \rangle} \hat{p} \rightsquigarrow p : \tau \dashv \langle \mathcal{G}; Y \rangle$ *then* $\Delta \vdash p : \tau \dashv Y$.
2. *If* $\vdash^{\Delta; \langle \mathcal{A}; \Phi \rangle; b} \check{p} \rightsquigarrow p : \tau \dashv^{\langle \mathcal{G}; Y \rangle}$ *then* $\Delta \vdash p : \tau \dashv Y$.

Proof. By mutual rule induction on Rules (4.6) and Rules (4.9).

1. We induct on the premise. In the following, let $\hat{Y} = \langle \mathcal{G}; Y \rangle$ and $\hat{\Phi} = \langle \mathcal{A}; \Phi \rangle$.

   Case (4.6a). We have:

$$
\begin{array}{lll}
(1) & \hat{p} = \hat{x} & \text{by assumption} \\
(2) & p = x & \text{by assumption} \\
(3) & Y = x : \tau & \text{by assumption} \\
(4) & \Delta \vdash x : \tau \dashv x : \tau & \text{by Rule (4.3a)}
\end{array}
$$

   Case (4.6b). We have:

$$
\begin{array}{lll}
(1) & p = \mathtt{wildp} & \text{by assumption} \\
(2) & Y = \varnothing & \text{by assumption} \\
(3) & \Delta \vdash \mathtt{wildp} : \tau \dashv \varnothing & \text{by Rule (4.3b)}
\end{array}
$$

Case (4.6c). We have:

(1) $\hat{p} = \mathtt{ufoldp}(\hat{p}')$ by assumption
(2) $p = \mathtt{foldp}(p')$ by assumption
(3) $\tau = \mathtt{rec}(t.\tau')$ by assumption
(4) $\Delta \vdash_{\hat{\Phi}} \hat{p}' \leadsto p' : [\mathtt{rec}(t.\tau')/t]\tau' \dashv\!\vdash \hat{Y}$ by assumption
(5) $\Delta \vdash p' : [\mathtt{rec}(t.\tau')/t]\tau' \dashv\!\vdash Y$ by IH, part 1 on (4)
(6) $\Delta \vdash \mathtt{foldp}(p') : \mathtt{rec}(t.\tau') \dashv\!\vdash Y$ by Rule (4.3c) on (5)

Case (4.6d). We have:

(1) $\hat{p} = \mathtt{utplp}[L](\{i \hookrightarrow \hat{p}_i\}_{i \in L})$ by assumption
(2) $p = \mathtt{tplp}[L](\{i \hookrightarrow p_i\}_{i \in L})$ by assumption
(3) $\tau = \mathtt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L})$ by assumption
(4) $\{\Delta \vdash_{\hat{\Phi}} \hat{p}_i \leadsto p_i : \tau_i \dashv\!\vdash \langle \mathcal{G}_i; Y_i \rangle\}_{i \in L}$ by assumption
(5) $Y = \cup_{i \in L} Y_i$ by assumption
(6) $\{\Delta \vdash p_i : \tau_i \dashv\!\vdash Y_i\}_{i \in L}$ by IH, part 1 over (4)
(7) $\Delta \vdash \mathtt{tplp}[L](\{i \hookrightarrow p_i\}_{i \in L}) : \mathtt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L}) \dashv\!\vdash \cup_{i \in L} Y_i$
     by Rule (4.3d) on (6)

Case (4.6e). We have:

(1) $\hat{p} = \mathtt{uinp}[\ell](\hat{p}')$ by assumption
(2) $p = \mathtt{inp}[\ell](p')$ by assumption
(3) $\tau = \mathtt{sum}[L, \ell](\{i \hookrightarrow \tau_i\}_{i \in L}; \ell \hookrightarrow \tau')$ by assumption
(4) $\Delta \vdash_{\hat{\Phi}} \hat{p}' \leadsto p' : \tau' \dashv\!\vdash \hat{Y}$ by assumption
(5) $\Delta \vdash p' : \tau' \dashv\!\vdash Y$ by IH, part 1 on (4)
(6) $\Delta \vdash \mathtt{inp}[\ell](p') : \mathtt{sum}[L, \ell](\{i \hookrightarrow \tau_i\}_{i \in L}; \ell \hookrightarrow \tau') \dashv\!\vdash Y$
     by Rule (4.3e) on (5)

Case (4.6f). We have:

(1) $\hat{p} = \mathtt{uapuptsm}[b][\hat{a}]$ by assumption
(2) $\mathcal{A} = \mathcal{A}', \hat{a} \leadsto a$ by assumption
(3) $\Phi = \Phi', a \hookrightarrow \mathtt{uptsm}(\tau; e_{\mathrm{parse}})$ by assumption
(4) $b \downarrow e_{\mathrm{body}}$ by assumption
(5) $e_{\mathrm{parse}}(e_{\mathrm{body}}) \Downarrow \mathtt{inj\ Success} \cdot e_{\mathrm{cand}}$ by assumption
(6) $e_{\mathrm{cand}} \uparrow_{\mathrm{CEPat}} \grave{p}$ by assumption
(7) $\vdash^{\Delta; \langle \mathcal{A}', \hat{a} \leadsto a; \Phi, a \hookrightarrow \mathtt{uptsm}(\tau; e_{\mathrm{parse}}) \rangle; b} \grave{p} \leadsto p : \tau \dashv\!\vdash^{\langle \mathcal{G}; Y \rangle}$ by assumption
(8) $\Delta \vdash p : \tau \dashv\!\vdash Y$ by IH, part 2 on (7)

2. We induct on the premise. In the following, let $\hat{Y} = \langle \mathcal{G}; Y \rangle$ and $\hat{\Phi} = \langle \mathcal{A}; \Phi \rangle$.

Case (4.9a). We have:

(1) $p = \mathtt{wildp}$ by assumption
(2) $Y = \varnothing$ by assumption
(3) $\Delta \vdash \mathtt{wildp} : \tau \dashv\!\vdash \varnothing$ by Rule (4.3b)

Case (4.9b). We have:

(1) $\grave{p} = \mathtt{cefoldp}(\grave{p}')$ by assumption

$\quad$ (2) $p = \text{foldp}(p')$ $\hfill$ by assumption

$\quad$ (3) $\tau = \text{rec}(t.\tau')$ $\hfill$ by assumption

$\quad$ (4) $\Delta \vdash \Phi \text{ upTSMs}$ $\hfill$ by assumption

$\quad$ (5) $\vdash^{\Delta;\hat{\Phi};b} \grave{p}' \rightsquigarrow p' : [\text{rec}(t.\tau')/t]\tau' \dashv\vert^{\hat{Y}}$ $\hfill$ by assumption

$\quad$ (6) $\Delta \vdash p' : [\text{rec}(t.\tau')/t]\tau' \dashv\vert Y$ $\hfill$ by IH, part 2 on (5)

$\hfill$ and (4)

$\quad$ (7) $\Delta \vdash \text{foldp}(p') : \text{rec}(t.\tau') \dashv\vert Y$ $\hfill$ by Rule (4.3c) on (6)

Case (4.9c). We have:

$\quad$ (1) $\grave{p} = \text{cetplp}[L](\{i \hookrightarrow \grave{p}_i\}_{i\in L})$ $\hfill$ by assumption

$\quad$ (2) $p = \text{tplp}[L](\{i \hookrightarrow p_i\}_{i\in L})$ $\hfill$ by assumption

$\quad$ (3) $\tau = \text{prod}[L](\{i \hookrightarrow \tau_i\}_{i\in L})$ $\hfill$ by assumption

$\quad$ (4) $\{\vdash^{\Delta;\hat{\Phi};b} \grave{p}_i \rightsquigarrow p_i : \tau_i \dashv\vert^{\langle \mathcal{G}_i; Y_i\rangle}\}_{i\in L}$ $\hfill$ by assumption

$\quad$ (5) $Y = \cup_{i\in L} Y_i$ $\hfill$ by assumption

$\quad$ (6) $\{\Delta \vdash p_i : \tau_i \dashv\vert Y_i\}_{i\in L}$ $\hfill$ by IH, part 2 over (4)

$\quad$ (7) $\Delta \vdash \text{tplp}[L](\{i \hookrightarrow p_i\}_{i\in L}) : \text{prod}[L](\{i \hookrightarrow \tau_i\}_{i\in L}) \dashv\vert \cup_{i\in L} Y_i$

$\hfill$ by Rule (4.3d) on (6)

Case (4.9d). We have:

$\quad$ (1) $\grave{p} = \text{ceinp}[\ell](\grave{p}')$ $\hfill$ by assumption

$\quad$ (2) $p = \text{inp}[\ell](p')$ $\hfill$ by assumption

$\quad$ (3) $\tau = \text{sum}[L,\ell](\{i \hookrightarrow \tau_i\}_{i\in L}; \ell \hookrightarrow \tau')$ $\hfill$ by assumption

$\quad$ (4) $\vdash^{\Delta;\hat{\Phi};b} \grave{p}' \rightsquigarrow p' : \tau' \dashv\vert^{\hat{Y}}$ $\hfill$ by assumption

$\quad$ (5) $\Delta \vdash p' : \tau' \dashv\vert Y$ $\hfill$ by IH, part 2 on (4)

$\quad$ (6) $\Delta \vdash \text{inp}[\ell](p') : \text{sum}[L,\ell](\{i \hookrightarrow \tau_i\}_{i\in L}; \ell \hookrightarrow \tau') \dashv\vert Y$

$\hfill$ by Rule (4.3e) on (5)

Case (4.9e). We have:

$\quad$ (1) $\grave{p} = \text{cesplicedp}[m;n]$ $\hfill$ by assumption

$\quad$ (2) $\text{parseUExp}(\text{subseq}(b;m;n)) = \hat{p}$ $\hfill$ by assumption

$\quad$ (3) $\Delta \vdash_{\hat{\Phi}} \hat{p} \rightsquigarrow p : \tau \dashv\vert \hat{Y}$ $\hfill$ by assumption

$\quad$ (4) $\Delta \vdash p : \tau \dashv\vert Y$ $\hfill$ by IH, part 1 on (3)

The mutual induction can be shown to be well-founded by showing that the following numeric metric on the judgements that we induct on is decreasing:

$$\|\Delta \vdash_{\hat{\Phi}} \hat{p} \rightsquigarrow p : \tau \dashv\vert \hat{Y}\| = \|\hat{p}\|$$

$$\|\vdash^{\Delta;\hat{\Phi};b} \grave{p} \rightsquigarrow p : \tau \dashv\vert^{\hat{Y}}\| = \|b\|$$

where $\|b\|$ is the length of b and $\|\hat{p}\|$ is the sum of the lengths of the literal bodies in $\hat{p}$,

$$\|\hat{x}\| = 0$$

$$\|\text{ufoldp}(\hat{p})\| = \|\hat{p}\|$$

$$\|\text{utplp}[L](\{i \hookrightarrow \hat{p}_i\}_{i\in L})\| = \sum_{i\in L} \|\hat{p}_i\|$$

$$\|\texttt{uinp}[\ell](\hat{p})\| = \|\hat{p}\|$$
$$\|\texttt{uapuptsm}[b][\hat{a}]\| = \|b\|$$

The only case in the proof of part 1 that invokes part 2 is Case (4.6f). There, we have that the metric remains stable:

$$\|\Delta \vdash_{\hat{\Phi}, \hat{a} \rightsquigarrow a \hookrightarrow \texttt{uetsm}(\tau; e_{\text{parse}})} \texttt{uapuptsm}[b][\hat{a}] \rightsquigarrow p : \tau \dashv \| \hat{Y} \|$$
$$= \|\vdash^{\Delta; \hat{\Phi}, \hat{a} \rightsquigarrow a \hookrightarrow \texttt{uetsm}(\tau; e_{\text{parse}}); b} \hat{p} \rightsquigarrow p : \tau \dashv \|^{\hat{Y}} \|$$
$$= \|b\|$$

The only case in the proof of part 2 that invokes part 1 is Case (4.9e). There, we have that $\text{parseUPat}(\text{subseq}(b; m; n)) = \hat{p}$ and the IH is applied to the judgement $\Delta \vdash_{\hat{\Phi}} \hat{p} \rightsquigarrow p : \tau \dashv \hat{Y}$. Because the metric is stable when passing from part 1 to part 2, we must have that it is strictly decreasing in the other direction:

$$\|\Delta \vdash_{\hat{\Phi}} \hat{p} \rightsquigarrow p : \tau \dashv \|\hat{Y}\| < \|\vdash^{\Delta; \hat{\Phi}; b} \texttt{cesplicedp}[m; n] \rightsquigarrow p : \tau \dashv \|^{\hat{Y}} \|$$

i.e. by the definitions above,

$$\|\hat{p}\| < \|b\|$$

This is established by appeal to Condition 3.22, which states that subsequences of b are no longer than b, and the following condition, which states that an unexpanded pattern constructed by parsing a textual sequence b is strictly smaller, as measured by the metric defined above, than the length of b, because some characters must necessarily be used to apply the pattern TSM and delimit each literal body.

Condition 4.22 (Pattern Parsing Monotonicity). *If* $\text{parseUPat}(b) = \hat{p}$ *then* $\|\hat{p}\| < \|b\|$.

Combining Conditions 3.22 and 4.22, we have that $\|\hat{e}\| < \|b\|$ as needed. $\qquad \square$

Finally, the following theorem establishes that typed expression and rule expansion produces expanded expressions and rules of the same type under the same contexts. Again, it must be stated mutually with the corresponding theorem about candidate expansion expressions and rules because the judgements are mutually defined.

Theorem 4.23 (Typed Expansion). *All of the following hold:*

1. (a) *If* $\langle \mathcal{D}; \Delta \rangle \langle \mathcal{G}; \Gamma \rangle \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{e} \rightsquigarrow e : \tau$ *then* $\Delta \, \Gamma \vdash e : \tau$.
    (b) *If* $\langle \mathcal{D}; \Delta \rangle \langle \mathcal{G}; \Gamma \rangle \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{r} \rightsquigarrow r : \tau \Mapsto \tau'$ *then* $\Delta \, \Gamma \vdash r : \tau \Mapsto \tau'$.

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

Proof. By mutual rule induction on Rules (4.4), Rule (4.5), Rules (4.7) and Rule (4.8).

1. In the following, let $\hat{\Delta} = \langle \mathcal{D}; \Delta \rangle$ and $\hat{\Gamma} = \langle \mathcal{G}; \Gamma \rangle$.

(a) We induct on the premise.

Case (4.4a) through (4.4j). These cases follow like the corresponding cases in the proof of Theorem 3.21, i.e. we apply Lemma 3.12 to or over the type expansion premises and the IH, part 1(a), to or over the typed expression expansion premises and then apply the corresponding typing rule in Rules (4.1a) through (4.1j).

Case (4.4k). We have:

(1)	$\hat{e} = \mathtt{umatch}[n]\{\hat{\tau}\}(\hat{e}'; \{\hat{r}_i\}_{1 \le i \le n})$	by assumption
(2)	$e = \mathtt{match}[n]\{\tau\}(e'; \{r_i\}_{1 \le i \le n})$	by assumption
(3)	$\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e}' \rightsquigarrow e' : \tau'$	by assumption
(4)	$\hat{\Delta} \vdash \hat{\tau} \rightsquigarrow \tau\ \text{type}$	by assumption
(5)	$\{\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{r}_i \rightsquigarrow r_i : \tau' \Mapsto \tau\}_{1 \le i \le n}$	by assumption
(6)	$\Delta\,\Gamma \vdash e' : \tau'$	by IH, part 1(a) on (3)
(7)	$\Delta \vdash \tau\ \text{type}$	by Lemma 3.12 on (4)
(8)	$\{\Delta\,\Gamma \vdash r_i : \tau' \Mapsto \tau\}_{1 \le i \le n}$	by IH, part 1(b) over (5)
(9)	$\Delta\,\Gamma \vdash \mathtt{match}[n]\{\tau\}(e'; \{r_i\}_{1 \le i \le n}) : \tau$	by Rule (4.1k) on (6), (7) and (8)

Case (4.4l). We have:

(1)	$\hat{e} = \mathtt{usyntaxue}\{e_{\text{parse}}\}\{\hat{\tau}'\}(\hat{a}.\hat{e}')$	by assumption
(2)	$\hat{\Delta} \vdash \hat{\tau}' \rightsquigarrow \tau'\ \text{type}$	by assumption
(3)	$\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi},\hat{a} \rightsquigarrow a \hookrightarrow \mathtt{uetsm}(\tau'; e_{\text{parse}});\hat{\Phi}} \hat{e}' \rightsquigarrow e : \tau$	by assumption
(4)	$\Delta\,\Gamma \vdash e : \tau$	by IH, part 1(a) on (3)

Case (4.4m). We have:

(1)	$\hat{e} = \mathtt{uapuetsm}[b][\hat{a}]$	by assumption
(2)	$\hat{\Psi} = \hat{\Psi}', \hat{a} \rightsquigarrow a \hookrightarrow \mathtt{uetsm}(\tau; e_{\text{parse}})$	by assumption
(3)	$b \downarrow e_{\text{body}}$	by assumption
(4)	$e_{\text{parse}}(e_{\text{body}}) \Downarrow \mathtt{inj\ Success} \cdot e_{\text{cand}}$	by assumption
(5)	$e_{\text{cand}} \uparrow_{\mathsf{CEExp}} \grave{e}$	by assumption
(6)	$\varnothing\,\varnothing \vdash^{\hat{\Delta};\hat{\Gamma};\hat{\Psi};\hat{\Phi};b} \grave{e} \rightsquigarrow e : \tau$	by assumption
(7)	$\varnothing \cap \Delta = \varnothing$	by finite set intersection identity
(8)	$\varnothing \cap \mathrm{dom}(\Gamma) = \varnothing$	by finite set intersection identity
(9)	$\varnothing \cup \Delta\,\varnothing \cup \Gamma \vdash e : \tau$	by IH, part 2(a) on (6), (7), and (8)
(10)	$\Delta\,\Gamma \vdash e : \tau$	by definition of finite set and finite function union over (9)

Case (4.4n). We have:

(1) $\hat{e} = \texttt{usyntaxup}\{e_{\text{parse}}\}\{\hat{\tau}'\}(\hat{a}.\hat{e}')$ by assumption

(2) $\hat{\Delta} \vdash \hat{\tau}' \leadsto \tau'$ type   by assumption

(3) $\hat{\Delta}\,\hat{\Gamma} \vdash_{\Psi;\hat{\Phi},\hat{a} \leadsto a \hookrightarrow \texttt{uptsm}(\tau';e_{\text{parse}})} \hat{e}' \leadsto e : \tau$ by assumption

(4) $\Delta\,\Gamma \vdash e : \tau$ by IH, part 1(a) on (3)

(b) We induct on the premise. There is only one case.

Case (4.5). We have:

(1) $\hat{r} = \texttt{urule}(\hat{p}.\hat{e})$ by assumption

(2) $r = \texttt{rule}(p.e)$ by assumption

(3) $\Delta \vdash_{\hat{\Phi}} \hat{p} \leadsto p : \tau \dashv \langle \mathcal{A}'; Y \rangle$ by assumption

(4) $\hat{\Delta}\,\langle \mathcal{A} \uplus \mathcal{A}'; \Gamma \cup Y \rangle \vdash_{\Psi;\hat{\Phi}} \hat{e} \leadsto e : \tau'$ by assumption

(5) $\Delta \vdash p : \tau \dashv Y$ by Theorem 4.21, part 1 on (3)

(6) $\Delta\,\Gamma \cup Y \vdash e : \tau'$ by IH, part 1(a) on (4)

(7) $\Delta\,\Gamma \vdash \texttt{rule}(p.e) : \tau \Mapsto \tau'$ by Rule (4.2) on (5) and (6)

2. In the following, let $\hat{\Delta} = \langle \mathcal{D}; \Delta_{\text{app}} \rangle$ and $\hat{\Gamma} = \langle \mathcal{G}; \Gamma_{\text{app}} \rangle$.

(a) We induct on the premise.

Case (4.7a) through (4.7j). These cases follow like the analagous cases in the proof of Theorem 3.21, i.e. we apply the IH, part 2(a) to all ce-expression validation judgements, Lemma 3.19 to all ce-type validation judgements, the identification convention to ensure that extended contexts remain disjoint, weakening and exchange as needed, and conclude by applying the corresponding typing rule in Rules (4.1a) through (4.1j).

Case (4.7k). We have:

(1) $\grave{e} = \texttt{cematch}[n]\{\grave{\tau}\}(\grave{e}'; \{\grave{r}_i\}_{1 \leq i \leq n})$ by assumption

(2) $e = \texttt{match}[n]\{\tau\}(e'; \{r_i\}_{1 \leq i \leq n})$ by assumption

(3) $\Delta\,\Gamma \vdash^{\hat{\Delta};\hat{\Gamma};\hat{\Psi};\hat{\Phi};b} \grave{e}' \leadsto e' : \tau'$ by assumption

(4) $\Delta \vdash^{\hat{\Delta};b} \grave{\tau} \leadsto \tau$ type   by assumption

(5) $\{\Delta\,\Gamma \vdash^{\hat{\Delta};\hat{\Gamma};\hat{\Psi};\hat{\Phi};b} \grave{r}_i \leadsto r_i : \tau' \Mapsto \tau\}_{1 \leq i \leq n}$ by assumption

(6) $\Delta \cap \Delta_{\text{app}} = \varnothing$ by assumption

(7) $\text{dom}(\Gamma) \cap \text{dom}(\Gamma_{\text{app}}) = \varnothing$ by assumption

(8) $\Delta \cup \Delta_{\text{app}}\,\Gamma \cup \Gamma_{\text{app}} \vdash e' : \tau'$ by IH, part 2(a) on (3), (6) and (7)

(9) $\Delta \cup \Delta_{\text{app}} \vdash \tau$ type   by Lemma 3.19 on (4)

(10) $\Delta \cup \Delta_{\text{app}}\,\Gamma \cup \Gamma_{\text{app}} \vdash r : \tau' \Mapsto \tau$ by IH, part 2(b) on (5), (6) and (7)

(11) $\Delta \cup \Delta_{\text{app}}\,\Gamma \cup \Gamma_{\text{app}} \vdash \texttt{match}[n]\{\tau\}(e'; \{r_i\}_{1 \leq i \leq n}) : \tau$ by Rule (4.1k) on (8), (9), (10)

Case (4.7l). We have:

(1) $\grave{e} = \texttt{cesplicede}[m;n]$ by assumption

$\quad$ (2) $\mathsf{parseUExp}(\mathsf{subseq}(b;m;n)) = \hat{e}$ $\qquad$ by assumption

$\quad$ (3) $\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \rightsquigarrow e : \tau$ $\qquad$ by assumption

$\quad$ (4) $\Delta \cap \Delta_{\mathsf{app}} = \varnothing$ $\qquad$ by assumption

$\quad$ (5) $\mathsf{dom}(\Gamma) \cap \mathsf{dom}(\Gamma_{\mathsf{app}}) = \varnothing$ $\qquad$ by assumption

$\quad$ (6) $\Delta_{\mathsf{app}}\,\Gamma_{\mathsf{app}} \vdash e : \tau$ $\qquad$ by IH, part 1(a) on (3)

$\quad$ (7) $\Delta \cup \Delta_{\mathsf{app}}\,\Gamma \cup \Gamma_{\mathsf{app}} \vdash e : \tau$ $\qquad$ by Lemma 4.1 over Δ and Γ and exchange on (6)

(b) We induct on the premise. There is only one case.

$\quad$ **Case** (4.8)**.** We have:

$\qquad$ (1) $\grave{r} = \mathtt{cerule}(p.\grave{e})$ $\qquad$ by assumption

$\qquad$ (2) $r = \mathtt{rule}(p.e)$ $\qquad$ by assumption

$\qquad$ (3) $\Delta \vdash p : \tau \dashv Y$ $\qquad$ by assumption

$\qquad$ (4) $\Delta\,\Gamma \cup Y \vdash^{\hat{\Delta};\hat{\Gamma};\hat{\Psi};\hat{\Phi};b} \grave{e} \rightsquigarrow e : \tau'$ $\qquad$ by assumption

$\qquad$ (5) $\Delta \cap \Delta_{\mathsf{app}} = \varnothing$ $\qquad$ by assumption

$\qquad$ (6) $\mathsf{dom}(\Gamma) \cap \mathsf{dom}(Y) = \varnothing$ $\qquad$ by identification convention

$\qquad$ (7) $\mathsf{dom}(\Gamma_{\mathsf{app}}) \cap \mathsf{dom}(Y) = \varnothing$ $\qquad$ by identification convention

$\qquad$ (8) $\mathsf{dom}(\Gamma) \cap \mathsf{dom}(\Gamma_{\mathsf{app}}) = \varnothing$ $\qquad$ by assumption

$\qquad$ (9) $\mathsf{dom}(\Gamma \cup Y) \cap \mathsf{dom}(\Gamma_{\mathsf{app}}) = \varnothing$ $\qquad$ by standard finite set definitions and identities on (6), (7) and (8)

$\qquad$ (10) $\Delta \cup \Delta_{\mathsf{app}}\,\Gamma \cup Y \cup \Gamma_{\mathsf{app}} \vdash e : \tau'$ $\qquad$ by IH, part 2(a) on (4), (5) and (9)

$\qquad$ (11) $\Delta \cup \Delta_{\mathsf{app}}\,\Gamma \cup \Gamma_{\mathsf{app}} \cup Y \vdash e : \tau'$ $\qquad$ by exchange of Y and Γ_{app} on (10)

$\qquad$ (12) $\Delta \cup \Delta_{\mathsf{app}}\,\Gamma \cup \Gamma_{\mathsf{app}} \vdash \mathtt{rule}(p.e) : \tau \mapsto \tau'$ $\qquad$ by Rule (4.2) on (3) and (11)

The mutual induction can be shown to be well-founded by showing that the following numeric metric on the judgements that we induct on is decreasing:

$$\|\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \rightsquigarrow e : \tau\| = \|\hat{e}\|$$

$$\|\Delta\,\Gamma \vdash^{\hat{\Delta};\hat{\Gamma};\hat{\Psi};\hat{\Phi};b} \grave{e} \rightsquigarrow e : \tau\| = \|b\|$$

where $\|b\|$ is the length of b and $\|\hat{e}\|$ is the sum of the lengths of the ueTSM literal bodies

in $\hat{e}$,

$$\|\hat{x}\| = 0$$
$$\|\texttt{ulam}\{\hat{\tau}\}(\hat{x}.\hat{e})\| = \|\hat{e}\|$$
$$\|\texttt{uap}(\hat{e}_1; \hat{e}_2)\| = \|\hat{e}_1\| + \|\hat{e}_2\|$$
$$\|\texttt{utlam}(\hat{t}.\hat{e})\| = \|\hat{e}\|$$
$$\|\texttt{utap}\{\hat{\tau}\}(\hat{e})\| = \|\hat{e}\|$$
$$\|\texttt{ufold}\{\hat{t}.\hat{\tau}\}(\hat{e})\| = \|\hat{e}\|$$
$$\|\texttt{uunfold}(\hat{e})\| = \|\hat{e}\|$$
$$\|\texttt{utpl}[L](\{i \hookrightarrow \hat{e}_i\}_{i \in L})\| = \sum_{i \in L} \|\hat{e}_i\|$$
$$\|\texttt{upr}[\ell](\hat{e})\| = \|\hat{e}\|$$
$$\|\texttt{uin}[L][\ell]\{\{i \hookrightarrow \hat{\tau}_i\}_{i \in L}\}(\hat{e})\| = \|\hat{e}\|$$
$$\|\texttt{umatch}[n]\{\hat{\tau}\}(\hat{e}; \{\hat{r}_i\}_{1 \le i \le n})\| = \|\hat{e}\| + \sum_{1 \le i \le n} \|r_i\|$$
$$\|\texttt{usyntaxue}\{e_{\text{parse}}\}\{\hat{\tau}\}(\hat{a}.\hat{e})\| = \|\hat{e}\|$$
$$\|\texttt{uapuetsm}[b][\hat{a}]\| = \|b\|$$
$$\|\texttt{usyntaxup}\{e_{\text{parse}}\}\{\hat{\tau}\}(\hat{a}.\hat{e})\| = \|\hat{e}\|$$

and $\|r\|$ is defined as follows:

$$\|\texttt{urule}(\hat{p}.\hat{e})\| = \|\hat{e}\|$$

The only case in the proof of part 1 that invokes part 2 is Case (4.4m). There, we have that the metric remains stable:

$$\|\hat{\Delta}\, \hat{\Gamma} \vdash_{\Psi, \hat{a} \leadsto a \hookrightarrow \texttt{uetsm}(\tau; e_{\text{parse}}); \hat{\Phi}} \texttt{uapuetsm}[b][\hat{a}] \leadsto e : \tau\|$$
$$= \|\varnothing\, \varnothing \vdash^{\hat{\Delta}; \hat{\Gamma}; \Psi, \hat{a} \leadsto a \hookrightarrow \texttt{uetsm}(\tau; e_{\text{parse}}); \hat{\Phi}; b} \hat{e} \leadsto e : \tau\|$$
$$= \|b\|$$

The only case in the proof of part 2 that invokes part 1 is Case (4.7l). There, we have that $\mathsf{parseUExp}(\mathsf{subseq}(b; m; n)) = \hat{e}$ and the IH is applied to the judgement $\hat{\Delta}\, \hat{\Gamma} \vdash_{\Psi; \hat{\Phi}} \hat{e} \leadsto e : \tau$. Because the metric is stable when passing from part 1 to part 2, we must have that it is strictly decreasing in the other direction:

$$\|\hat{\Delta}\, \hat{\Gamma} \vdash_{\Psi; \hat{\Phi}} \hat{e} \leadsto e : \tau\| < \|\Delta\, \Gamma \vdash^{\hat{\Delta}; \hat{\Gamma}; \Psi; \hat{\Phi}; b} \texttt{cesplicede}[m; n] \leadsto e : \tau\|$$

i.e. by the definitions above,

$$\|\hat{e}\| < \|b\|$$

This is established by appeal to Condition 3.22, which states that subsequences of b are no longer than b, and the following condition, which states that an unexpanded expression constructed by parsing a textual sequence b is strictly smaller, as measured by the metric defined above, than the length of b, because some characters must necessarily be used to apply a TSM and delimit each literal body.

Condition 4.24 (Expression Parsing Monotonicity)**.** *If* parseUExp$(b) = \hat{e}$ *then* $\|\hat{e}\| < \|b\|$.

Combining Conditions 3.22 and 4.24, we have that $\|\hat{e}\| < \|b\|$ as needed. $\qquad\square$

Part II

Parametric TSMs

Chapter 5

Parameterized TSMs (pTSMs)

You know me, I gotta put in a big tree.

— Bob Ross, *The Joy of Painting*

In the preceding chapters, we introduced unparameterized TSMs (uTSMs). uTSMs are defined at a single type, like Rx, and the expansions that they generate have access to bindings at the application site only via spliced subterms. In this chapter, we introduce *parameterized TSMs* (pTSMs) – TSMs parameterized by types and modules. pTSMs can be defined over a parameterized family of types, and can access their parameters within the expansions that they generate.

This chapter is organized much like the preceding chapters. We begin in Sec. 5.1 by introducing parameterized TSMs by example in VerseML. In particular, we discuss type parameters in Sec. 5.1.1 and module parameters in Sec. 5.1.2. We then formalize parameterized TSMs by defining a reduced calculus, miniVerse$_\forall$, in Sec. 5.2.

5.1 Parameterized TSMs By Example

5.1.1 Type Parameters

VerseML, like many ML dialects, provides support for defining *type constructors*, which express type-parameterized families of types. The canonical example is the type constructor list, which constructs list types given one type parameter, the *element type*. In VerseML, list is defined as follows.

```
type list('a) = Nil | Cons of 'a * list('a)
```

For any type T, the type of lists containing elements of type T can be constructed by applying list to T, written list(T).[1] In other words, type constructors can be understood as total functions at the level of *constructor expressions*.

---

[1]As with function application, the parentheses are optional (though for type constructor application, it is typical to include them). In many other ML dialects, type parameters are given in prefix form, e.g. in Standard ML, one writes `int list` rather than `list(int)`.

Kinds classify constructor expressions, much like types classify expressions. Types are constructor expressions of kind `T`, and type constructors are constructor expressions of arrow kind. Here, `list` takes a single type parameter, so it has arrow kind `T -> T`.

As discussed in Sec. **??**, full-scale ML dialects commonly define derived syntactic forms that decrease the syntactic cost of introducing and pattern matching over values of list type. VerseML, in contrast, does not build in such derived list forms.

In lieu of derived forms for introducing lists, we define the following *parameterized expression TSM* (peTSM):

```
1  syntax $list('a) at list('a) {
2    static fn(body : Body) : ParseResult(CEExp) => (* ... *)
3  }
```

Line 1 names the peTSM `$list` and specifies a single type parameter, `'a` (implicitly of kind `T`). This type parameter appears in the type annotation, which specifies that `$list`, when applied to a type `'a` and a generalized literal form, will only generate expansions of type `list('a)`. For example, we can apply `$list` to the type `int` and generalized literal forms delimited by square brackets as follows.

```
val y = $list(int) [3, 4, 5]
val x = $list(int) [1, 2 :: y]
```

Line 2 of the definition of `$list` defines its parse function. Parse functions operate as described in Chapter 3 to generate encodings of candidate expansions, which are subsequently validated to generate the final expansions of expressions of peTSM application form, like those in the example above. For `$list`, the parse function (whose body is elided above for concision) breaks the literal body up into spliced subexpressions – those separated by commas become individual elements at the head of the list being generated, and, optionally, a trailing spliced subexpression prefixed by two colons (`::`) becomes the tail of the list being generated (the tail is `Nil` otherwise). The final expansion of the example above, if written textually, is:

```
val y : list(int) = Cons(3, Cons(4, Cons(5, Nil)))
val x : list(int) = Cons(1, Cons(2, y))
```

Similarly, in lieu of derived list pattern forms, we define the following *parameterized pattern TSM* (ppTSM):

```
syntax $list('a) at list('a) for patterns {
  static fn(body : Body) : ParseResult(CEPat) => (* ... *)
}
```

Again, Line 1 names the ppTSM `$list` and specifies a single type parameter, `'a`. This type parameter appears in the type annotation, which specifies that `$list`, when apply to a type `'a` and a generalized literal form, will only generate patterns that match values of type `list('a)`.

For example, we can apply the ppTSM `$list` and the `$list` to define the polymorphic map function as follows.

```
fun map (f : 'a -> 'b) (x : list('a)) => match x {
  $list('a) [] => $list('b) []
  | $list('a) [hd :: tl] => $list('b) [f hd :: map f tl]
```

}

The expansion of this function definition, written textually, is:

```
fun map (f : 'a -> 'b) (x : list('a)) : 'b list => match x {
  Nil => Nil
| Cons(hd, tl) => Cons(f hd, map f tl)
}
```

This is somewhat unsatisfying, however, because the expansion is more concise than the unexpanded definition of map. To further reduce syntactic cost, we can designate $list as the implicit TSM for both expressions and patterns at all types 'a list around our definition of map as follows.

```
implicit syntax $list('a) in
  fun map (f : 'a -> 'b) (x : 'a list) : 'b list => match x {
    [] => []
  | [hd :: tl] => [f hd :: map f tl]
  }
end
```

By designating an implicit TSM, we no longer need to explicitly apply $list within expressions in analytic position or patterns.

5.1.2 Module Parameters

VerseML also provides a module language based on the Standard ML module language [44]. The module language consists of *module expressions* classified by *signatures*.

The canonical example is the signature for working with persistent queues.

```
signature QUEUE = sig
  type queue('a)
  val empty  : queue('a)
  val insert : 'a * queue('a) -> queue('a)
  val remove : queue('a) -> option('a * queue('a))
end
```

Structures that match this signature must define a type constructor queue of kind T -> T and three values – empty introduces the empty queue, insert inserts a value onto the back of a queue, and remove removes the element at the front of the queue and returns it and the remaining queue, or None if the queue is empty.

There are many possible structures that implement this signature. For example, we can define a structure ListQueue that represents queues internally as lists, where the head of the list is the back of the queue. With this representation, insert is a constant time operation, but remove is a linear time operation. Alternatively, we might define a structure TwoListQueue that represents queues internally as a pair of lists, maintaining the invariant that one is the reverse of the other, so that both insert and remove are constant time operations (see [27] for the details of this and other possibilities).

Regardless of the implementation that the client chooses, we would like for the client to be able to introduce queues more naturally and at lower syntactic cost than is possible by directly applying the functions specified by the signature above. In VerseML, we

can give clients of structures matching the signature QUEUE this ability by defining the following parameterized expression TSM:

```
syntax $queue(Q : QUEUE)('a) at Q.queue('a) {
  static fn(body : Body) : ParseResult(CEExp) => (* ... *)
}
```

This peTSM specifies one module parameter, Q, which must match the signature QUEUE, and one type parameter, 'a (implicitly of kind T). These appear in the type annotation, which specifies that expansions that arise from applying $queue to a module Q : QUEUE and a type 'a will be of type Q.queue('a). For example:

```
1 val q = $queue TwoListQueue int [> 1, 2, 3]
2 val q' = $queue TwoListQueue int [q > 4, 5]
```

On Line 1, the initial angle bracket (>) indicates that the items are inserted in left-to-right order. The items in the queue are given as spliced subexpressions separated by commas. Line 2 inserts two additional items onto the back of the queue q. The expansion of this example, written textually, is:

```
1 val q : TwoListQueue.queue(int) =
2   TwoListQueue.insert(1,
3     TwoListQueue.insert(2,
4       TwoListQueue.insert(3,
5         TwoListQueue.empty)))
6 val q' : TwoListQueue.queue(int) =
7   TwoListQueue.insert(4, TwoListQueue.insert(5, q))
```

Notice that the expansion can refer to the module parameter TwoListQueue.

We can further reduce syntactic cost by defining a synonym for the partial application of $queue to the module parameter TwoListQueue:

```
syntax $tlq = $queue TwoListQueue
val q = $tlq int [> 1, 2, 3]
```

We can further define a synonym for the partial application of $tlq to a type parameter:

```
syntax $tlqi = $tlq int (* = $queue TwoListQueue int *)
val q' = $tlqi [q > 4, 5]
```

Another way to reduce syntactic cost is by designating $queue Q 'a the implicit TSM at all types of the form Q.queue('a) where Q : QUEUE. This is written as follows:

```
implicit syntax (Q : QUEUE) ('a) => $queue Q 'a in
  val q : TwoListQueue.queue(int) = [> 1, 2, 3]
  val q' : TwoListQueue.queue(int) = [q > 4, 5]
end
```

This designation is particularly useful for clients who need to construct a queue as an argument to a function. For example, consider a function

```
enqueue_jobs : Q.queue(Job) -> Ticket
```

for some module Q : QUEUE and types Job and Ticket. We can enqueue a sequence of jobs j1 through j4 under the TSM designation above as follows:

```
enqueue_jobs [> j1, j2, j3, j4]
```

Sort			Operational Form	Stylized Form	Description
Kind	κ	$::=$	$\mathtt{darr}(\kappa;u.\kappa)$	$(u :: \kappa) \to \kappa$	dependent function
			$\mathtt{unit}$	$\langle\!\langle\,\rangle\!\rangle$	nullary product
			$\mathtt{dprod}(\kappa;u.\kappa)$	$(u :: \kappa) \times \kappa$	dependent product
			$\mathtt{Type}$	$\mathtt{T}$	types
			$\mathtt{S}(\tau)$	$[\!=\!\tau]$	singleton
Con	c,τ	$::=$	u	u	variable
			t	t	variable
			$\mathtt{abs}(u.c)$	$\lambda u.c$	abstraction
			$\mathtt{app}(c;c)$	$c(c)$	application
			$\mathtt{triv}$	$\langle\!\langle\,\rangle\!\rangle$	trivial
			$\mathtt{pair}(c;c)$	$\langle\!\langle c,c \rangle\!\rangle$	pair
			$\mathtt{prl}(c)$	$c \cdot \mathtt{l}$	left projection
			$\mathtt{prr}(c)$	$c \cdot \mathtt{r}$	right projection
			$\mathtt{parr}(\tau;\tau)$	$\tau \rightharpoonup \tau$	partial function
			$\mathtt{all}\{\kappa\}(u.\tau)$	$\forall (u :: \kappa).\tau$	polymorphic
			$\mathtt{rec}(t.\tau)$	$\mu t.\tau$	recursive
			$\mathtt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L})$	$\langle \{i \hookrightarrow \tau_i\}_{i \in L} \rangle$	labeled product
			$\mathtt{sum}[L](\{i \hookrightarrow \tau_i\}_{i \in L})$	$[\{i \hookrightarrow \tau_i\}_{i \in L}]$	labeled sum
			$\mathtt{con}(M)$	$M \cdot \mathtt{c}$	constructor part

Figure 5.1: Syntax of kinds and constructors in miniVerse$_\forall$. By convention, we choose the metavariable τ for constructors that, in well-formed terms, must necessarily be of kind $\mathtt{T}$, and the metavariable c otherwise. Similarly, we use constructor variables t to stand for constructors of kind $\mathtt{T}$, and constructor variables u otherwise. Kinds and constructors are identified up to α-equivalence.

Sort			Operational Form	Stylized Form	Description
Exp	e	$::=$	x	x	variable
			$\mathtt{lam}\{\tau\}(x.e)$	$\lambda x{:}\tau.e$	abstraction
			$\mathtt{ap}(e;e)$	$e(e)$	application
			$\mathtt{clam}\{\kappa\}(u.e)$	$\Lambda u{::}\kappa.e$	constructor abstraction
			$\mathtt{cap}\{\kappa\}(e)$	$e[\kappa]$	constructor application
			$\mathtt{fold}\{t.\tau\}(e)$	$\mathtt{fold}(e)$	fold
			$\mathtt{unfold}(e)$	$\mathtt{unfold}(e)$	unfold
			$\mathtt{tpl}[L](\{i \hookrightarrow e_i\}_{i \in L})$	$\langle\{i \hookrightarrow e_i\}_{i \in L}\rangle$	labeled tuple
			$\mathtt{pr}[\ell](e)$	$e \cdot \ell$	projection
			$\mathtt{in}[L][\ell]\{\{i \hookrightarrow \tau_i\}_{i \in L}\}(e)$	$\mathtt{inj}\ \ell \cdot e$	injection
			$\mathtt{match}[n]\{\tau\}(e; \{r_i\}_{1 \leq i \leq n})$	$\mathtt{match}\ e\ \{r_i\}_{1 \leq i \leq n}$	match
			$\mathtt{val}(M)$	$M \cdot \mathbf{v}$	value part
Rule	r	$::=$	$\mathtt{rule}(p.e)$	$p \Rightarrow e$	rule
Pat	p	$::=$	x	x	variable pattern
			$\mathtt{wildp}$	$\_$	wildcard pattern
			$\mathtt{foldp}(p)$	$\mathtt{fold}(p)$	fold pattern
			$\mathtt{tplp}[L](\{i \hookrightarrow p_i\}_{i \in L})$	$\langle\{i \hookrightarrow p_i\}_{i \in L}\rangle$	labeled tuple pattern
			$\mathtt{inp}[\ell](p)$	$\mathtt{inj}\ \ell \cdot p$	injection pattern

Figure 5.2: Syntax of expanded expressions, rules and patterns (collectively, expanded terms) in miniVerse$_\forall$. Expanded terms are identified up to α-equivalence.

Sort			Operational Form	Stylized Form	Description
Sig	σ	$::=$	$\mathtt{sig}\{\kappa\}(u.\tau)$	$[\![u :: \kappa; \tau]\!]$	signature
Mod	M	$::=$	X	X	variable
			$\mathtt{struct}(c;e)$	$[\![c;e]\!]$	structure
			$\mathtt{seal}\{\sigma\}(M)$	$M \uparrow \sigma$	seal
			$\mathtt{mlet}\{\sigma\}(M;X.M)$	$(\mathtt{let}\ X = M\ \mathtt{in}\ M) : \sigma$	definition

Figure 5.3: Syntax of signatures and module expressions in miniVerse$_\forall$. Signatures and module expressions are identified up to α-equivalence.

5.2 miniVerse$_\forall$

5.2.1 Syntax of the Inner Language

Syntax of the Inner Core Language

We adopt the metatheoretic conventions established for our definitions of miniVerse$_{\mathbf{UE}}$ in Sec. 3.2 and miniVerse$_{\mathbf{U}}$ in Sec. 4.2 without restating them.

Kinds and constructors in Figure 5.1.

Expanded expressions, rules and patterns in Figure 5.2.

Syntax of the Inner Module Language

Module expressions and signatures in Figure 5.3.

5.2.2 Statics of the Inner Language

The *statics of the inner language* is defined by a collection of judgements that we organize into three groups.

The first group of judgements, which we refer to as the *statics of the inner constructor language*, define the statics of expanded kinds and constructors.

Judgement Form	Description
$\Omega \vdash \kappa$ kind	κ is a well-formed kind
$\Omega \vdash \kappa \equiv \kappa'$	κ and κ' are equivalent
$\Omega \vdash \kappa <:: \kappa'$	κ is a subkind of κ'
$\Omega \vdash c :: \kappa$	c has kind κ
$\Omega \vdash c \equiv c' :: \kappa$	c and c' are equivalent as constructors of kind κ

The second group of judgements, which we refer to as the *statics of the inner core language*, define the statics of types, expanded expressions, rules and patterns.

Judgement Form	Description
$\Omega \vdash \tau$ type	τ is a well-formed type
$\Omega \vdash \tau \equiv \tau'$ type	τ and τ' are equivalent types
$\Omega \vdash \tau <: \tau'$	τ is a subtype of τ'
$\Omega \vdash e : \tau$	e is assigned type τ
$\Omega \vdash r : \tau \Mapsto \tau'$	r takes values of type τ to values of type τ'
$\Omega \vdash p : \tau \dashv \Omega'$	p matches values of type τ and generates hypotheses Ω'

The third group of judgements, which we refer to as the *statics of the inner module language*, define the statics of expanded signatures and module expressions.

Judgement Form   **Description**

$\Omega \vdash \sigma \; \mathsf{sig}$	σ is a well-formed signature
$\Omega \vdash \sigma \equiv \sigma'$	σ and σ' are definitionally equal
$\Omega \vdash \sigma <: \sigma'$	σ is a sub-signature of σ'
$\Omega \vdash M : \sigma$	M has signature σ
$\Omega \vdash M \; \mathsf{mval}$	M is a module value

A *unified inner context*, Ω, is a finite function that maps
- each expression variable $x \in \mathrm{dom}(\Omega)$ to the hypothesis $x : \tau$ for some τ;
- each constructor variable $u \in \mathrm{dom}(\Omega)$ to the hypothesis $u :: \kappa$ for some κ; and
- each module variable $X \in \mathrm{dom}(\Omega)$ to the hypothesis $X :: \sigma$ for some σ.

We write
- $\Omega, x : \tau$ when $x \notin \mathrm{dom}(\Omega)$ and $\Omega \vdash \tau :: \mathtt{Type}$ for the extension of Ω with a mapping from x to $x : \tau$
- $\Omega, u :: \kappa$ when $u \notin \mathrm{dom}(\Omega)$ and $\Omega \vdash \kappa \; \mathsf{kind}$ for the extension of Ω with a mapping from u to $u :: \kappa$
- $\Omega, X : \sigma$ when $X \notin \mathrm{dom}(\Omega)$ and $\Omega \vdash \sigma \; \mathsf{sig}$ for the extension of Ω with a mapping from X to $X : \sigma$.

A well-formed unified inner context is one that can be constructed by some sequence of such extensions, start from the empty context, $\varnothing$.

Kinds and Constructors

Kind formation TODO: describe these

$$\frac{\Omega \vdash \kappa_1 \; \mathsf{kind} \qquad \Omega, u :: \kappa_1 \vdash \kappa_2 \; \mathsf{kind}}{\Omega \vdash \mathtt{darr}(\kappa_1; u.\kappa_2) \; \mathsf{kind}} \tag{5.1a}$$

$$\frac{}{\Omega \vdash \mathtt{unit} \; \mathsf{kind}} \tag{5.1b}$$

$$\frac{\Omega \vdash \kappa_1 \; \mathsf{kind} \qquad \Omega, u :: \kappa_1 \vdash \kappa_2 \; \mathsf{kind}}{\Omega \vdash \mathtt{dprod}(\kappa_1; u.\kappa_2) \; \mathsf{kind}} \tag{5.1c}$$

$$\frac{}{\Omega \vdash \mathtt{Type} \; \mathsf{kind}} \tag{5.1d}$$

$$\frac{\Omega \vdash \tau :: \mathtt{Type}}{\Omega \vdash \mathtt{S}(\tau) \; \mathsf{kind}} \tag{5.1e}$$

Kind definitional equality TODO: describe these

$$\frac{\Omega \vdash \kappa \; \mathsf{kind}}{\Omega \vdash \kappa \equiv \kappa} \tag{5.2a}$$

$$\frac{\Omega \vdash \kappa \equiv \kappa'}{\Omega \vdash \kappa' \equiv \kappa} \tag{5.2b}$$

$$\frac{\Omega \vdash \kappa \equiv \kappa' \qquad \Omega \vdash \kappa' \equiv \kappa''}{\Omega \vdash \kappa \equiv \kappa''} \tag{5.2c}$$

$$\frac{\Omega \vdash \kappa_1 \equiv \kappa_1' \qquad \Omega, u :: \kappa_1 \vdash \kappa_2 \equiv \kappa_2'}{\Omega \vdash \mathtt{darr}(\kappa_1; u.\kappa_2) \equiv \mathtt{darr}(\kappa_1'; u.\kappa_2')} \tag{5.2d}$$

$$\frac{\Omega \vdash \kappa_1 \equiv \kappa_1' \qquad \Omega, u :: \kappa_1 \vdash \kappa_2 \equiv \kappa_2'}{\Omega \vdash \mathtt{dprod}(\kappa_1; u.\kappa_2) \equiv \mathtt{dprod}(\kappa_1'; u.\kappa_2')} \tag{5.2e}$$

$$\frac{\Omega \vdash c \equiv c' :: \mathtt{Type}}{\Omega \vdash \mathtt{S}(c) \equiv \mathtt{S}(c')} \tag{5.2f}$$

Subkinding TODO: describe these

$$\frac{\Omega \vdash \kappa \equiv \kappa'}{\Omega \vdash \kappa <:: \kappa'} \tag{5.3a}$$

$$\frac{\Omega \vdash \kappa <:: \kappa' \qquad \Omega \vdash \kappa' <:: \kappa''}{\Omega \vdash \kappa <:: \kappa''} \tag{5.3b}$$

$$\frac{\Omega \vdash \kappa_1' <:: \kappa_1 \qquad \Omega, u :: \kappa_1' \vdash \kappa_2 <:: \kappa_2'}{\Omega \vdash \mathtt{darr}(\kappa_1; u.\kappa_2) <:: \mathtt{darr}(\kappa_1'; u.\kappa_2')} \tag{5.3c}$$

$$\frac{\Omega \vdash \kappa_1 <:: \kappa_1' \qquad \Omega, u :: \kappa_1 \vdash \kappa_2 <:: \kappa_2'}{\Omega \vdash \mathtt{dprod}(\kappa_1; u.\kappa_2) <:: \mathtt{dprod}(\kappa_1'; u.\kappa_2')} \tag{5.3d}$$

$$\frac{\Omega \vdash \tau :: \mathtt{Type}}{\Omega \vdash \mathtt{S}(\tau) <:: \mathtt{Type}} \tag{5.3e}$$

$$\frac{\Omega \vdash \tau <: \tau'}{\Omega \vdash \mathtt{S}(\tau) <:: \mathtt{S}(\tau')\mathtt{Type}} \tag{5.3f}$$

Kinding TODO: describe these

$$\frac{\Omega \vdash c :: \kappa_1 \qquad \Omega \vdash \kappa_1 <:: \kappa_2}{\Omega \vdash c :: \kappa_2} \tag{5.4a}$$

$$\frac{}{\Omega, u :: \kappa \vdash u :: \kappa} \tag{5.4b}$$

$$\frac{\Omega, u :: \kappa_1 \vdash c_2 :: \kappa_2}{\Omega \vdash \mathtt{abs}(u.c_2) :: \mathtt{darr}(\kappa_1; u.\kappa_2)} \tag{5.4c}$$

$$\frac{\Omega \vdash c_1 :: \mathtt{darr}(\kappa_2; u.\kappa) \qquad \Omega \vdash c_2 :: \kappa_2}{\Omega \vdash \mathtt{app}(c_1; c_2) :: [c_1/u]\kappa} \tag{5.4d}$$

$$\frac{}{\Omega \vdash \mathtt{triv} :: \mathtt{unit}} \tag{5.4e}$$

$$\frac{\Omega \vdash c_1 :: \kappa_1 \qquad \Omega \vdash c_2 :: [c_1/u]\kappa_2}{\Omega \vdash \mathtt{pair}(c_1; c_2) :: \mathtt{dprod}(\kappa_1; u.\kappa_2)} \tag{5.4f}$$

$$\frac{\Omega \vdash c :: \mathtt{dprod}(\kappa_1; u.\kappa_2)}{\Omega \vdash \mathtt{prl}(c) :: \kappa_1} \tag{5.4g}$$

$$\frac{\Omega \vdash c :: \mathtt{dprod}(\kappa_1; u.\kappa_2)}{\Omega \vdash \mathtt{prr}(c) :: [\mathtt{prl}(c)/u]\kappa_2} \tag{5.4h}$$

$$\frac{\Omega \vdash \tau_1 :: \mathtt{Type} \qquad \Omega \vdash \tau_2 :: \mathtt{Type}}{\Omega \vdash \mathtt{parr}(\tau_1; \tau_2) :: \mathtt{Type}} \tag{5.4i}$$

$$\frac{\Omega \vdash \kappa \ \mathtt{kind} \qquad \Omega, u :: \kappa \vdash \tau :: \mathtt{Type}}{\Omega \vdash \mathtt{all}\{\kappa\}(u.\tau) :: \mathtt{Type}} \tag{5.4j}$$

$$\frac{\Omega, t :: \mathtt{Type} \vdash \tau :: \mathtt{Type}}{\Omega \vdash \mathtt{rec}(t.\tau) :: \mathtt{Type}} \tag{5.4k}$$

$$\frac{\{\Omega \vdash \tau_i :: \mathtt{Type}\}_{1 \leq i \leq n}}{\Omega \vdash \mathtt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L}) :: \mathtt{Type}} \tag{5.4l}$$

$$\frac{\{\Omega \vdash \tau_i :: \mathtt{Type}\}_{1 \leq i \leq n}}{\Omega \vdash \mathtt{sum}[L](\{i \hookrightarrow \tau_i\}_{i \in L}) :: \mathtt{Type}} \tag{5.4m}$$

$$\frac{\Omega \vdash c :: \mathtt{Type}}{\Omega \vdash c :: \mathtt{S}(c)} \tag{5.4n}$$

$$\frac{\Omega \vdash M \ \mathtt{mval} \qquad \Omega \vdash M : \mathtt{sig}\{\kappa\}(u.\tau)}{\Omega \vdash \mathtt{con}(M) :: \kappa} \tag{5.4o}$$

Constructor equalityTODO: describe this

$$\frac{\Omega \vdash c :: \kappa}{\Omega \vdash c \equiv c :: \kappa} \tag{5.5a}$$

$$\frac{\Omega \vdash c \equiv c' :: \kappa}{\Omega \vdash c' \equiv c :: \kappa} \tag{5.5b}$$

$$\frac{\Omega \vdash c \equiv c' :: \kappa \qquad \Omega \vdash c' \equiv c'' :: \kappa}{\Omega \vdash c \equiv c'' :: \kappa} \tag{5.5c}$$

$$\frac{\Omega, u :: \kappa_1 \vdash c \equiv c' :: \kappa_2}{\Omega \vdash \mathtt{abs}(u.c) \equiv \mathtt{abs}(u.c') :: \mathtt{darr}(\kappa_1; u.\kappa_2)} \tag{5.5d}$$

$$\frac{\Omega \vdash c_1 \equiv c_1' :: \mathtt{darr}(\kappa_2; u.\kappa) \qquad \Omega \vdash c_2 \equiv c_2' :: \kappa_2}{\Omega \vdash \mathtt{app}(c_1; c_2) \equiv \mathtt{app}(c_1'; c_2') :: \kappa} \tag{5.5e}$$

$$\frac{\Omega \vdash \mathtt{abs}(u.c) :: \mathtt{darr}(\kappa_2; u.\kappa) \qquad \Omega \vdash c_2 :: \kappa_2}{\Omega \vdash \mathtt{app}(\mathtt{abs}(u.c); c_2) \equiv [c_2/u]c :: [c_2/u]\kappa} \tag{5.5f}$$

$$\frac{\Omega \vdash c_1 \equiv c_1' :: \kappa_1 \qquad \Omega \vdash c_2 \equiv c_2' :: [c_1/u]\kappa_2}{\Omega \vdash \mathtt{pair}(c_1; c_2) \equiv \mathtt{pair}(c_1'; c_2') :: \mathtt{dprod}(\kappa_1; u.\kappa_2)} \tag{5.5g}$$

$$\frac{\Omega \vdash c \equiv c' :: \mathtt{dprod}(\kappa_1; u.\kappa_2)}{\Omega \vdash \mathtt{prl}(c) \equiv \mathtt{prl}(c') :: \kappa_1} \tag{5.5h}$$

$$\frac{\Omega \vdash c_1 :: \kappa_1 \qquad \Omega \vdash c_2 :: \kappa_2}{\Omega \vdash \mathtt{prl}(\mathtt{pair}(c_1; c_2)) \equiv c_1 :: \kappa_1} \tag{5.5i}$$

$$\frac{\Omega \vdash c \equiv c' :: \mathtt{dprod}(\kappa_1; u.\kappa_2)}{\Omega \vdash \mathtt{prr}(c) \equiv \mathtt{prr}(c') :: [\mathtt{prl}(c)/u]\kappa_2} \tag{5.5j}$$

$$\frac{\Omega \vdash c_1 :: \kappa_1 \qquad \Omega \vdash c_2 :: \kappa_2}{\Omega \vdash \mathtt{prr}(\mathtt{pair}(c_1; c_2)) \equiv c_2 :: \kappa_2} \tag{5.5k}$$

$$\frac{\Omega \vdash \tau_1 \equiv \tau_1' :: \mathsf{Type} \qquad \Omega \vdash \tau_2 \equiv \tau_2' :: \mathsf{Type}}{\Omega \vdash \mathtt{parr}(\tau_1; \tau_2) \equiv \mathtt{parr}(\tau_1'; \tau_2') :: \mathsf{Type}} \tag{5.5l}$$

$$\frac{\Omega \vdash \kappa \equiv \kappa' \qquad \Omega, u :: \kappa \vdash \tau \equiv \tau' :: \mathsf{Type}}{\Omega \vdash \mathtt{all}\{\kappa\}(u.\tau) \equiv \mathtt{all}\{\kappa'\}(u.\tau') :: \mathsf{Type}} \tag{5.5m}$$

$$\frac{\Omega, t :: \mathsf{Type} \vdash \tau \equiv \tau' :: \mathsf{Type}}{\Omega \vdash \mathtt{rec}(t.\tau) \equiv \mathtt{rec}(t.\tau') :: \mathsf{Type}} \tag{5.5n}$$

$$\frac{\{\Omega \vdash \tau_i \equiv \tau_i' :: \mathsf{Type}\}_{1 \le i \le n}}{\Omega \vdash \mathtt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L}) \equiv \mathtt{prod}[L](\{i \hookrightarrow \tau_i'\}_{i \in L}) :: \mathsf{Type}} \tag{5.5o}$$

$$\frac{\{\Omega \vdash \tau_i \equiv \tau_i' :: \mathsf{Type}\}_{1 \le i \le n}}{\Omega \vdash \mathtt{sum}[L](\{i \hookrightarrow \tau_i\}_{i \in L}) \equiv \mathtt{sum}[L](\{i \hookrightarrow \tau_i'\}_{i \in L}) :: \mathsf{Type}} \tag{5.5p}$$

$$\frac{\Omega \vdash c :: \mathsf{S}(c')}{\Omega \vdash c \equiv c' :: \mathsf{Type}} \tag{5.5q}$$

$$\frac{\Omega \vdash \mathtt{struct}(c; e) \ \mathsf{mval} \qquad \Omega \vdash \mathtt{struct}(c; e) : \mathtt{sig}\{\kappa\}(u.\tau)}{\Omega \vdash \mathtt{con}(\mathtt{struct}(c; e)) \equiv c :: \kappa} \tag{5.5r}$$

Expanded Expressions, Rules and Patterns

Types, τ, classify expressions. The constructors of kind Type coincide with the types of miniVerse$_\forall$.

$$\frac{\Omega \vdash \tau :: \mathsf{Type}}{\Omega \vdash \tau \ \mathsf{type}} \tag{5.6}$$

Type equality then coincides with constructor equality at kind Type.

$$\frac{\Omega \vdash \tau \equiv \tau :: \mathtt{Type}}{\Omega \vdash \tau \equiv \tau' \; \mathtt{type}} \tag{5.7}$$

Subtyping.

$$\frac{\Omega \vdash \tau_1 \equiv \tau_2 \; \mathtt{type}}{\Omega \vdash \tau_1 <: \tau_2} \tag{5.8a}$$

$$\frac{\Omega \vdash \tau <: \tau' \qquad \Omega \vdash \tau' <: \tau''}{\Omega \vdash \tau <: \tau''} \tag{5.8b}$$

$$\frac{\Omega \vdash \tau_1' <: \tau_1 \qquad \Omega \vdash \tau_2 <: \tau_2'}{\Omega \vdash \mathtt{parr}(\tau_1; \tau_2) <: \mathtt{parr}(\tau_1'; \tau_2')} \tag{5.8c}$$

$$\frac{\Omega \vdash \kappa <:: \kappa' \qquad \Omega, u :: \kappa \vdash \tau <: \tau'}{\Omega \vdash \mathtt{all}\{\kappa\}(u.\tau) <: \mathtt{all}\{\kappa'\}(u.\tau')} \tag{5.8d}$$

$$\frac{\{\Omega \vdash \tau_i <: \tau_i'\}_{i \in L}}{\Omega \vdash \mathtt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L}) <: \mathtt{prod}[L](\{i \hookrightarrow \tau_i'\}_{i \in L})} \tag{5.8e}$$

$$\frac{\{\Omega \vdash \tau_i <: \tau_i'\}_{i \in L}}{\Omega \vdash \mathtt{sum}[L](\{i \hookrightarrow \tau_i\}_{i \in L}) <: \mathtt{sum}[L](\{i \hookrightarrow \tau_i'\}_{i \in L})} \tag{5.8f}$$

Expression typing TODO: describe these

$$\frac{\Omega \vdash e : \tau \qquad \Omega \vdash \tau <: \tau'}{\Omega \vdash e : \tau'} \tag{5.9a}$$

$$\frac{}{\Omega, x : \tau \vdash x : \tau} \tag{5.9b}$$

$$\frac{\Omega \vdash \tau \; \mathtt{type} \qquad \Omega, x : \tau \vdash e : \tau'}{\Omega \vdash \mathtt{lam}\{\tau\}(x.e) : \mathtt{parr}(\tau; \tau')} \tag{5.9c}$$

$$\frac{\Omega \vdash e_1 : \mathtt{parr}(\tau; \tau') \qquad \Omega \vdash e_2 : \tau}{\Omega \vdash \mathtt{ap}(e_1; e_2) : \tau'} \tag{5.9d}$$

$$\frac{\Omega \vdash \kappa \; \mathtt{kind} \qquad \Omega, u :: \kappa \vdash e : \tau}{\Omega \vdash \mathtt{clam}\{\kappa\}(u.e) : \mathtt{all}\{\kappa\}(u.\tau)} \tag{5.9e}$$

$$\frac{\Omega \vdash e : \mathtt{all}\{\kappa\}(u.\tau) \qquad \Omega \vdash c :: \kappa}{\Omega \vdash \mathtt{cap}\{c\}(e) : [c/u]\tau} \tag{5.9f}$$

$$\frac{\Omega, t :: \mathtt{Type} \vdash \tau \; \mathtt{type} \qquad \Omega \vdash e : [\mathtt{rec}(t.\tau)/t]\tau}{\Omega \vdash \mathtt{fold}\{t.\tau\}(e) : \mathtt{rec}(t.\tau)} \tag{5.9g}$$

$$\frac{\Omega \vdash e : \mathtt{rec}(t.\tau)}{\Omega \vdash \mathtt{unfold}(e) : [\mathtt{rec}(t.\tau)/t]\tau} \tag{5.9h}$$

$$\frac{\{\Omega \vdash e_i : \tau_i\}_{i \in L}}{\Omega \vdash \mathtt{tpl}[L](\{i \hookrightarrow e_i\}_{i \in L}) : \mathtt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L})} \tag{5.9i}$$

$$\frac{\Omega \vdash e : \mathtt{prod}[L, \ell](\{i \hookrightarrow \tau_i\}_{i \in L}; \ell \hookrightarrow \tau)}{\Omega \vdash \mathtt{pr}[\ell](e) : \tau} \tag{5.9j}$$

$$\frac{\{\Omega \vdash \tau_i \,\mathsf{type}\}_{i \in L} \qquad \Omega \vdash \tau \,\mathsf{type} \qquad \Omega \vdash e : \tau}{\Omega \vdash \mathtt{in}[L, \ell][\ell]\{\{i \hookrightarrow \tau_i\}_{i \in L}; \ell \hookrightarrow \tau\}(e) : \mathtt{sum}[L, \ell](\{i \hookrightarrow \tau_i\}_{i \in L}; \ell \hookrightarrow \tau)} \tag{5.9k}$$

$$\frac{\Omega \vdash e : \tau \qquad \Omega \vdash \tau' \,\mathsf{type} \qquad \{\Omega \vdash r_i : \tau \Mapsto \tau'\}_{1 \leq i \leq n}}{\Omega \vdash \mathtt{match}[n]\{\tau'\}(e; \{r_i\}_{1 \leq i \leq n}) : \tau'} \tag{5.9l}$$

$$\frac{\Omega \vdash M \,\mathsf{mval} \qquad \Omega \vdash M : \mathtt{sig}\{\kappa\}(u.\tau)}{\Omega \vdash \mathtt{val}(M) : [\mathtt{con}(M)/u]\tau} \tag{5.9m}$$

maybe last one needs self-recognition?

Rule typing

$$\frac{\Omega \vdash p : \tau \dashv\!\mid \Omega' \qquad \Omega \cup \Omega' \vdash e : \tau'}{\Omega \vdash \mathtt{rule}(p.e) : \tau \Mapsto \tau'} \tag{5.10}$$

Pattern typing

$$\frac{\Omega \vdash p : \tau \dashv\!\mid \Omega' \qquad \Omega \vdash \tau <: \tau'}{\Omega \vdash p : \tau' \dashv\!\mid \Omega'} \tag{5.11a}$$

$$\frac{}{\Omega \vdash x : \tau \dashv\!\mid x : \tau} \tag{5.11b}$$

$$\frac{}{\Omega \vdash \mathtt{wildp} : \tau \dashv\!\mid \varnothing} \tag{5.11c}$$

$$\frac{\Omega \vdash p : [\mathtt{rec}(t.\tau)/t]\tau \dashv\!\mid \Omega'}{\Omega \vdash \mathtt{foldp}(p) : \mathtt{rec}(t.\tau) \dashv\!\mid \Omega'} \tag{5.11d}$$

$$\frac{\{\Omega \vdash p_i : \tau_i \dashv\!\mid \Omega_i\}_{i \in L}}{\Omega \vdash \mathtt{tplp}[L](\{i \hookrightarrow p_i\}_{i \in L}) : \mathtt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L}) \dashv\!\mid \bigcup_{i \in L} \Omega_i} \tag{5.11e}$$

$$\frac{\Omega \vdash p : \tau \dashv\!\mid \Omega'}{\Omega \vdash \mathtt{inp}[\ell](p) : \mathtt{sum}[L, \ell](\{i \hookrightarrow \tau_i\}_{i \in L}; \ell \hookrightarrow \tau) \dashv\!\mid \Omega'} \tag{5.11f}$$

Signatures and Structures

Signature formation

$$\frac{\Omega \vdash \kappa \,\mathsf{kind} \qquad \Omega, u :: \kappa \vdash \tau \,\mathsf{type}}{\Omega \vdash \mathtt{sig}\{\kappa\}(u.\tau) \,\mathsf{sig}} \tag{5.12}$$

Signature equality

$$\frac{\Omega \vdash \kappa \equiv \kappa' \qquad \Omega, u :: \kappa \vdash \tau \equiv \tau' \,\mathsf{type}}{\Omega \vdash \mathtt{sig}\{\kappa\}(u.\tau) \equiv \mathtt{sig}\{\kappa'\}(u.\tau')} \tag{5.13}$$

Subsignature

$$\frac{\Omega \vdash \kappa <:: \kappa' \qquad \Omega, u :: \kappa \vdash \tau <: \tau'}{\Omega \vdash \mathtt{sig}\{\kappa\}(u.\tau) <: \mathtt{sig}\{\kappa'\}(u.\tau')} \tag{5.14}$$

Signature matching

$$\frac{\Omega \vdash M : \sigma \qquad \Omega \vdash \sigma <: \sigma'}{\Omega \vdash M : \sigma'} \tag{5.15a}$$

$$\frac{}{\Omega, X : \sigma \vdash X : \sigma} \tag{5.15b}$$

$$\frac{\Omega \vdash c :: \kappa \qquad \Omega \vdash e : [c/u]\tau}{\Omega \vdash \mathtt{struct}(c;e) : \mathtt{sig}\{\kappa\}(u.\tau)} \tag{5.15c}$$

$$\frac{\Omega \vdash \sigma \ \mathsf{sig} \qquad \Omega \vdash M : \sigma}{\Omega \vdash \mathtt{seal}\{\sigma\}(M) : \sigma} \tag{5.15d}$$

$$\frac{\Omega \vdash M : \sigma \qquad \Omega \vdash \sigma' \ \mathsf{sig} \qquad \Omega, X : \sigma \vdash M' : \sigma'}{\Omega \vdash \mathtt{mlet}\{\sigma'\}(M; X.M') : \sigma'} \tag{5.15e}$$

Module values

$$\frac{}{\Omega \vdash \mathtt{struct}(c;e) \ \mathsf{mval}} \tag{5.16a}$$

$$\frac{}{\Omega, X : \sigma \vdash X \ \mathsf{mval}} \tag{5.16b}$$

TODO: metatheory of statics of inner core

5.2.3 Structural Dynamics

Expressions

Judgement Form	Description
$e \mapsto e'$	e transitions to e'
$\Omega \vdash e \ \mathsf{val}$	e is a value

Modules

Judgement Form	Description
$M \mapsto M'$	M transitions to M'
$\Omega \vdash M \ \mathsf{mval}$	M is a module value

5.2.4 Syntax of the Surface Language

Syntax of the Surface Core Language

Unexpanded kinds and constructors in Figure 5.4
Unexpanded expressions, rules and patterns in Figure 5.5

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Sort			Operational Form	Stylized Form	Description
UKind	$\hat{\kappa}$	$::=$	$\texttt{udarr}(\hat{\kappa};\hat{u}.\hat{\kappa})$	$(\hat{u} :: \hat{\kappa}) \to \hat{\kappa}$	dependent function
			$\texttt{uunit}$	$\langle\!\langle\,\rangle\!\rangle$	nullary product
			$\texttt{udprod}(\hat{\kappa};\hat{u}.\hat{\kappa})$	$(\hat{u} :: \hat{\kappa}) \times \hat{\kappa}$	dependent product
			$\texttt{uType}$	T	types
			$\texttt{uS}(\hat{\tau})$	$[=\hat{\tau}]$	singleton
UCon	$\hat{c}, \hat{\tau}$	$::=$	$\hat{u}$	$\hat{u}$	constructor sigil
			$\hat{t}$	$\hat{t}$	constructor sigil
			$\texttt{uasc}\{\hat{\kappa}\}(\hat{c})$	$\hat{c} :: \hat{\kappa}$	ascription
			$\texttt{uabs}(\hat{u}.\hat{c})$	$\lambda\hat{u}.\hat{c}$	abstraction
			$\texttt{uapp}(c;c)$	$c(c)$	application
			$\texttt{utriv}$	$\langle\!\langle\,\rangle\!\rangle$	trivial
			$\texttt{upair}(\hat{c};\hat{c})$	$\langle\!\langle \hat{c}, \hat{c} \rangle\!\rangle$	pair
			$\texttt{uprl}(\hat{c})$	$\hat{c} \cdot \mathtt{l}$	left projection
			$\texttt{uprr}(\hat{c})$	$\hat{c} \cdot \mathtt{r}$	right projection
			$\texttt{uparr}(\hat{\tau};\hat{\tau})$	$\hat{\tau} \rightharpoonup \hat{\tau}$	partial function
			$\texttt{uall}\{\hat{\kappa}\}(\hat{u}.\hat{\tau})$	$\forall(\hat{u} :: \hat{\kappa}).\hat{\tau}$	polymorphic
			$\texttt{urec}(\hat{t}.\hat{\tau})$	$\mu\hat{t}.\hat{\tau}$	recursive
			$\texttt{uprod}[L](\{i \hookrightarrow \hat{\tau}_i\}_{i \in L})$	$\langle \{i \hookrightarrow \hat{\tau}_i\}_{i \in L} \rangle$	labeled product
			$\texttt{usum}[L](\{i \hookrightarrow \hat{\tau}_i\}_{i \in L})$	$[\{i \hookrightarrow \hat{\tau}_i\}_{i \in L}]$	labeled sum
			$\texttt{ucon}(\hat{X})$	$\hat{X} \cdot \mathtt{c}$	constructor part

Figure 5.4: Syntax of unexpanded kinds and constructors in miniVerse$_\forall$. By convention, we choose the metavariable $\hat{\tau}$ for constructors that, in well-formed terms, must necessarily expand to constructors of kind T, and the metavariable $\hat{c}$ otherwise. Similarly, we choose metavariables $\hat{t}$ for constructor sigils that expand to constructor variables that stand for constructors of kind T, and constructor sigils $\hat{u}$ otherwise. Unexpanded kinds and constructors are not identified up to α-equivalence.

DRAFT (March 30, 2016)

Sort			Operational Form	Stylized Form	Description
UExp	$\hat{e}$	::=	$\hat{x}$	$\hat{x}$	sigil
			$\mathtt{uasc}\{\hat{\tau}\}(\hat{e})$	$\hat{e} : \hat{\tau}$	ascription
			$\mathtt{uletval}(\hat{e}; \hat{x}.\hat{e})$	$\mathtt{let\,val}\ \hat{x} = \hat{e}\ \mathtt{in}\ \hat{e}$	value binding
			$\mathtt{uanalam}(\hat{x}.\hat{e})$	$\lambda\hat{x}.\hat{e}$	abstraction (unannotated)
			$\mathtt{ulam}\{\hat{\tau}\}(\hat{x}.\hat{e})$	$\lambda\hat{x}{:}\hat{\tau}.\hat{e}$	abstraction (annotated)
			$\mathtt{uap}(\hat{e}; \hat{e})$	$\hat{e}(\hat{e})$	application
			$\mathtt{uclam}\{\hat{\kappa}\}(\hat{u}.\hat{e})$	$\Lambda\hat{u}{::}\hat{\kappa}.\hat{e}$	constructor abstraction
			$\mathtt{ucap}\{\hat{c}\}(\hat{e})$	$\hat{e}[\hat{c}]$	constructor application
			$\mathtt{ufold}(\hat{e})$	$\mathtt{fold}(\hat{e})$	fold
			$\mathtt{uunfold}(\hat{e})$	$\mathtt{unfold}(\hat{e})$	unfold
			$\mathtt{utpl}[L](\{i \hookrightarrow \hat{e}_i\}_{i \in L})$	$\langle\{i \hookrightarrow \hat{e}_i\}_{i \in L}\rangle$	labeled tuple
			$\mathtt{upr}[\ell](\hat{e})$	$\hat{e} \cdot \ell$	projection
			$\mathtt{uin}[\ell](\hat{e})$	$\mathtt{inj}\ \ell \cdot \hat{e}$	injection
			$\mathtt{umatch}[n](\hat{e}; \{\hat{r}_i\}_{1 \leq i \leq n})$	$\mathtt{match}\ \hat{e}\ \{\hat{r}_i\}_{1 \leq i \leq n}$	match
			$\mathtt{uval}(\hat{X})$	$\hat{X} \cdot \mathbf{v}$	value part
			$\mathtt{usyntaxpe}\{e\}\{\hat{\rho}\}(\hat{a}.\hat{e})$	$\mathtt{syntax}\ \hat{a}\ \mathtt{at}\ \hat{\rho}\ \mathtt{for}$ $\mathtt{expressions}\ \{e\}\ \mathtt{in}\ \hat{e}$	peTSM definition
			$\mathtt{uletpetsm}\{\hat{e}\}(\hat{a}.\hat{e})$	$\mathtt{let\,syntax}\ \hat{a} = \hat{e}\ \mathtt{for}$ $\mathtt{expressions\ in}\ \hat{e}$	peTSM binding
			...	...	peTSM designation
			$\mathtt{uappetsm}[b]\{\hat{e}\}$	$\hat{e}\ /b/$	peTSM application
			$\mathtt{uelit}[b]$	$/b/$	peTSM unadorned literal
			$\mathtt{usyntaxpp}\{e\}\{\hat{\rho}\}(\hat{a}.\hat{e})$	$\mathtt{syntax}\ \hat{a}\ \mathtt{at}\ \hat{\rho}\ \mathtt{for}$ $\mathtt{patterns}\ \{e\}\ \mathtt{in}\ \hat{e}$	ppTSM definition
			$\mathtt{uletpptsm}\{\hat{e}\}(\hat{a}.\hat{e})$	$\mathtt{let\,syntax}\ \hat{a} = \hat{e}\ \mathtt{for}$ $\mathtt{patterns\ in}\ \hat{e}$	ppTSM binding
			...	...	ppTSM designation
URule	$\hat{r}$	::=	$\mathtt{urule}(\hat{p}.\hat{e})$	$\hat{p} \Rightarrow \hat{e}$	match rule
UPat	$\hat{p}$	::=	$\hat{x}$	$\hat{x}$	sigil pattern
			$\mathtt{uwildp}$	$\_$	wildcard pattern
			$\mathtt{ufoldp}(\hat{p})$	$\mathtt{fold}(\hat{p})$	fold pattern
			$\mathtt{utplp}[L](\{i \hookrightarrow \hat{p}_i\}_{i \in L})$	$\langle\{i \hookrightarrow \hat{p}_i\}_{i \in L}\rangle$	labeled tuple pattern
			$\mathtt{uinp}[\ell](\hat{p})$	$\mathtt{inj}\ \ell \cdot \hat{p}$	injection pattern
			$\mathtt{uappptsm}[b]\{\hat{e}\}$	$\hat{e}\ /b/$	ppTSM application
			$\mathtt{uplit}[b]$	$/b/$	ppTSM unadorned literal

Figure 5.5: Abstract syntax of unexpanded expressions, rules and patterns in miniVerse$_\forall$.

Sort			Operational Form	Stylized Form	Description
USig	$\hat{\sigma}$	$::=$	$\mathtt{usig}\{\hat{\kappa}\}(\hat{u}.\hat{\tau})$	$[\![\hat{u} :: \hat{\kappa}; \hat{\tau}]\!]$	signature
UMod	$\hat{M}$	$::=$	$\hat{X}$	$\hat{X}$	module sigil
			$\mathtt{ustruct}(\hat{c};\hat{e})$	$[\![\hat{c};\hat{e}]\!]$	structure
			$\mathtt{useal}\{\hat{\sigma}\}(\hat{M})$	$\hat{M} \upharpoonright \hat{\sigma}$	seal
			$\mathtt{umlet}\{\hat{\sigma}\}(\hat{M}; \hat{X}.\hat{M})$	$(\mathtt{let}\ \hat{X} = \hat{M}\ \mathtt{in}\ \hat{M}) : \hat{\sigma}$	definition
			$\mathtt{umsyntaxpe}\{e\}\{\hat{\rho}\}(\hat{a}.\hat{M})$	syntax $\hat{a}$ at $\hat{\rho}$ for expressions $\{e\}$ in $\hat{M}$	peTSM definition
			$\mathtt{umletpetsm}\{\hat{e}\}(\hat{a}.\hat{M})$	let syntax $\hat{a} = \hat{e}$ for expressions in $\hat{M}$	peTSM binding
			...	...	peTSM designation
			$\mathtt{usyntaxpp}\{e\}\{\hat{\rho}\}(\hat{a}.\hat{M})$	syntax $\hat{a}$ at $\hat{\rho}$ for patterns $\{e\}$ in $\hat{M}$	ppTSM definition
			$\mathtt{uletpptsm}\{\hat{e}\}(\hat{a}.\hat{M})$	let syntax $\hat{a} = \hat{e}$ for patterns in $\hat{M}$	ppTSM binding
			...	...	ppTSM designation

Figure 5.6: Abstract syntax of unexpanded module expressions and signatures in miniVerse$_\forall$.

Sort			Operational Form	Stylized Form	Description
MType	ρ	$::=$	$\mathtt{type}(\tau)$	τ	type annotation
			$\mathtt{alltypes}(t.\rho)$	$\forall t.\rho$	type abstraction
			$\mathtt{allmods}\{\sigma\}(X.\rho)$	$\forall X{:}\sigma.\rho$	module abstraction
UMType	$\hat{\rho}$	$::=$	$\mathtt{utype}(\hat{\tau})$	$\hat{\tau}$	type annotation
			$\mathtt{ualltypes}(\hat{t}.\hat{\rho})$	$\forall \hat{t}.\hat{\rho}$	type abstraction
			$\mathtt{uallmods}\{\hat{\sigma}\}(\hat{X}.\hat{\rho})$	$\forall \hat{X}{:}\hat{\sigma}.\hat{\rho}$	module abstraction
MExp	ϵ	$::=$	$\mathtt{defref}[a]$	a	TSM definition reference
			$\mathtt{abstype}(t.\epsilon)$	$\Lambda t.\epsilon$	type abstraction
			$\mathtt{absmod}\{\sigma\}(X.\epsilon)$	$\Lambda X{:}\sigma.\epsilon$	module abstraction
			$\mathtt{aptype}\{\tau\}(\epsilon)$	$\epsilon(\tau)$	type application
			$\mathtt{apmod}\{M\}(\epsilon)$	$\epsilon(M)$	module application
UMExp	$\hat{\epsilon}$	$::=$	$\mathtt{bindref}[\hat{a}]$	$\hat{a}$	TSM binding reference
			$\mathtt{uabstype}(\hat{t}.\hat{\epsilon})$	$\Lambda \hat{t}.\hat{\epsilon}$	type abstraction
			$\mathtt{uabsmod}\{\hat{\sigma}\}(\hat{X}.\hat{\epsilon})$	$\Lambda \hat{X}{:}\hat{\sigma}.\hat{\epsilon}$	module abstraction
			$\mathtt{uaptype}\{\hat{\tau}\}(\hat{\epsilon})$	$\hat{\epsilon}(\hat{\tau})$	type application
			$\mathtt{uapmod}\{\hat{M}\}(\hat{\epsilon})$	$\hat{\epsilon}(\hat{M})$	module application

Figure 5.7: Abstract syntax of TSM types, unexpanded TSM types, TSM expressions and unexpanded TSM expressions in miniVerse$_\forall$.

Syntax of the TSM Language

Macro types, unexpanded macro types, macro expressions and unexpanded macro expressions in Figure 5.7

Syntax of Type Patterns

TODO: do this

Syntax of the Surface Module Language

Figure 5.6

5.2.5 Typed Expansion

Unexpanded kinds and constructors

Judgement Form	Description
$\hat{\Omega} \vdash \hat{\kappa} \rightsquigarrow \kappa \text{ kind}$	$\hat{\kappa}$ is well-formed and has expansion κ
$\hat{\Omega} \vdash \hat{c} \rightsquigarrow c \Rightarrow \kappa$	$\hat{c}$ has expansion c and synthesizes kind κ
$\hat{\Omega} \vdash \hat{c} \rightsquigarrow c \Leftarrow \kappa$	$\hat{c}$ has expansion c when analyzed against kind κ

Unexpanded types, expressions, rules and patterns

Judgement Form	Description
$\hat{\Omega} \vdash \hat{\tau} \rightsquigarrow \tau \text{ type}$	$\hat{\tau}$ has expansion τ
$\hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \rightsquigarrow e \Rightarrow \tau$	$\hat{e}$ has expansion e and synthesizes type τ
$\hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \rightsquigarrow e \Leftarrow \tau$	$\hat{e}$ has expansion e when analyzed against type τ
$\hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{r} \rightsquigarrow r \Rightarrow \tau \Mapsto \tau'$	$\hat{r}$ has expansion r and takes values of type τ to values of synthesized type τ'
$\hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{r} \rightsquigarrow r \Leftarrow \tau \Mapsto \tau'$	$\hat{r}$ has expansion r and takes values of type τ to values of type τ' when $\tau's$ is provided for analysis
$\hat{\Omega} \vdash_{\hat{\Phi}} \hat{p} \rightsquigarrow p : \tau \dashv\!\vdash \hat{\Omega}$	$\hat{p}$ has expansion p and type τ and generates hypotheses $\hat{\Omega}$

Unexpanded signatures and module expressions

Judgement Form	Description
$\hat{\Omega} \vdash \hat{\sigma} \rightsquigarrow \sigma \text{ sig}$	$\hat{\sigma}$ has expansion σ
$\hat{\Omega} \vdash_{\hat{\Psi}_{\text{E}};\hat{\Psi}_{\text{P}}} \hat{M} \rightsquigarrow M \Rightarrow \sigma$	$\hat{M}$ has expansion M and synthesizes signature σ
$\hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{M} \rightsquigarrow M \Leftarrow \sigma$	$\hat{M}$ has expansion M and matches signature σ

A *unified outer context*, $\hat{\Omega}$, takes the form $\langle \mathcal{D}; \mathcal{G}; \mathcal{M}; \Omega \rangle$, where $\mathcal{D}$ is a *constructor sigil expansion context*, $\mathcal{G}$ is an *expression sigil expansion context*, $\mathcal{M}$ is a *module sigil expansion context* and Ω is a unified inner context.

A constructor sigil expansion context, $\mathcal{D}$, is a finite function that maps each constructor sigil $\hat{u} \in \mathrm{dom}(\mathcal{D})$ to the constructor sigil expansion $\hat{u} \rightsquigarrow u$. We write $\hat{\Omega}, \hat{u} \rightsquigarrow u :: \kappa$ when $\hat{\Omega} = \langle \mathcal{D}; \mathcal{G}; \mathcal{M}; \Omega \rangle$ as an abbreviation of

$$\langle \mathcal{D} \uplus \hat{u} \rightsquigarrow u; \mathcal{G}; \mathcal{M}; \Omega, u :: \kappa \rangle$$

An expression sigil expansion context, $\mathcal{G}$, is a finite function that maps each expression sigil $\hat{x} \in \mathrm{dom}(\mathcal{G})$ to the expression sigil expansion $\hat{x} \rightsquigarrow x$. We write $\hat{\Omega}, \hat{x} \rightsquigarrow x : \tau$ when $\hat{\Omega} = \langle \mathcal{D}; \mathcal{G}; \mathcal{M}; \Omega \rangle$ as an abbreviation of

$$\langle \mathcal{D}; \mathcal{G} \uplus \hat{x} \rightsquigarrow x; \mathcal{M}; \Omega, x : \tau \rangle$$

A module sigil expansion context, $\mathcal{M}$, is a finite function that maps each module sigil $\hat{X} \in \mathrm{dom}(\mathcal{M})$ to the module sigil expansion $\hat{X} \rightsquigarrow X$. We write $\hat{\Omega}, \hat{X} \rightsquigarrow X : \sigma$ when $\hat{\Omega} = \langle \mathcal{D}; \mathcal{G}; \mathcal{M}; \Omega \rangle$ as an abbreviation of

$$\langle \mathcal{D}; \mathcal{G}; \mathcal{M} \uplus \hat{X} \rightsquigarrow X; \Omega, X : \sigma \rangle$$

A *peTSM context*, $\hat{\Psi}$, takes the form $\langle \mathcal{A}; \Psi; \mathcal{I} \rangle$ where $\mathcal{A}$ is a *TSM binding context*, Ψ is a *peTSM definition context*, and $\mathcal{I}$ is TODO: implicits. Similarly, a *ppTSM context*, $\hat{\Phi}$, takes the form $\langle \mathcal{A}; \Phi; \mathcal{I} \rangle$ where Φ is a *ppTSM definition context*.

A *TSM binding context*, $\mathcal{A}$, is a finite function that maps each TSM name $\hat{a} \in \mathrm{dom}(\mathcal{A})$ to a *TSM binding*, $\hat{a} \hookrightarrow \epsilon$, for some TSM expression, ϵ.

A *peTSM definition context*, Ψ, is a finite function that maps each symbol $a \in \mathrm{dom}(\Psi)$ to a *peTSM definition*, $a \hookrightarrow \mathtt{petsm}(\rho; e_{\mathrm{parse}})$ for some TSM type ρ and expanded expression e_{parse}.

Similarly, a *ppTSM definition context*, Φ, is a finite function that maps each symbol $a \in \mathrm{dom}(\Phi)$ to a *ppTSM definition*, $a \hookrightarrow \mathtt{pptsm}(\rho; e_{\mathrm{parse}})$ for some TSM type ρ and expanded expression e_{parse}.

Kinds and Constructors

Kind expansion

$$\frac{\hat{\Omega} \vdash \hat{\kappa}_1 \rightsquigarrow \kappa_1 \text{ kind} \qquad \hat{\Omega}, \hat{u} \rightsquigarrow u :: \kappa_1 \vdash \hat{\kappa}_2 \rightsquigarrow \kappa_2 \text{ kind}}{\hat{\Omega} \vdash \mathtt{udarr}(\hat{\kappa}_1; \hat{u}.\hat{\kappa}_2) \rightsquigarrow \mathtt{darr}(\kappa_1; u.\kappa_2) \text{ kind}} \tag{5.17a}$$

$$\frac{}{\hat{\Omega} \vdash \mathtt{uunit} \rightsquigarrow \mathtt{unit} \text{ kind}} \tag{5.17b}$$

$$\frac{\hat{\Omega} \vdash \hat{\kappa}_1 \rightsquigarrow \kappa_1 \text{ kind} \qquad \hat{\Omega}, \hat{u} \rightsquigarrow u :: \kappa_1 \vdash \hat{\kappa}_2 \rightsquigarrow \kappa_2 \text{ kind}}{\hat{\Omega} \vdash \mathtt{udprod}(\hat{\kappa}_1; \hat{u}.\hat{\kappa}_2) \rightsquigarrow \mathtt{dprod}(\kappa_1; u.\kappa_2) \text{ kind}} \tag{5.17c}$$

$$\frac{}{\hat{\Omega} \vdash \mathtt{uType} \rightsquigarrow \mathtt{Type} \text{ kind}} \tag{5.17d}$$

$$\frac{\hat{\Omega} \vdash \hat{\tau} \rightsquigarrow \tau \Leftarrow \mathtt{Type}}{\hat{\Omega} \vdash \mathtt{uS}(\hat{\tau}) \rightsquigarrow \mathtt{S}(\tau) \text{ kind}} \tag{5.17e}$$

Synthetic constructor expansion

$$\overline{\hat{\Omega}, \hat{u} \rightsquigarrow u :: \kappa \vdash \hat{u} \rightsquigarrow u \Rightarrow \kappa} \tag{5.18a}$$

$$\frac{\hat{\Omega} \vdash \hat{\kappa} \rightsquigarrow \kappa \text{ kind} \qquad \hat{\Omega} \vdash \hat{c} \rightsquigarrow c \Leftarrow \kappa}{\hat{\Omega} \vdash \texttt{uasc}\{\hat{\kappa}\}(\hat{c}) \rightsquigarrow c \Rightarrow \kappa} \tag{5.18b}$$

$$\frac{\hat{\Omega} \vdash \hat{c}_1 \rightsquigarrow c_1 \Rightarrow \texttt{darr}(\kappa_2; u.\kappa) \qquad \hat{\Omega} \vdash \hat{c}_2 \rightsquigarrow c_2 \Leftarrow \kappa_2}{\hat{\Omega} \vdash \texttt{uapp}(\hat{c}_1; \hat{c}_2) \rightsquigarrow \texttt{app}(c_1; c_2) \Rightarrow [c_1/u]\kappa} \tag{5.18c}$$

$$\overline{\hat{\Omega} \vdash \texttt{utriv} \rightsquigarrow \texttt{triv} \Rightarrow \texttt{unit}} \tag{5.18d}$$

$$\frac{\hat{\Omega} \vdash \hat{c} \rightsquigarrow c \Rightarrow \texttt{dprod}(\kappa_1; u.\kappa_2)}{\hat{\Omega} \vdash \texttt{uprl}(\hat{c}) \rightsquigarrow \texttt{prl}(c) \Rightarrow \kappa_1} \tag{5.18e}$$

$$\frac{\hat{\Omega} \vdash \hat{c} \rightsquigarrow c \Rightarrow \texttt{dprod}(\kappa_1; u.\kappa_2)}{\hat{\Omega} \vdash \texttt{uprr}(\hat{c}) \rightsquigarrow \texttt{prr}(c) \Rightarrow [\texttt{prl}(c)/u]\kappa_2} \tag{5.18f}$$

$$\frac{\hat{\Omega} \vdash \hat{\tau}_1 \rightsquigarrow \tau_1 \Leftarrow \texttt{Type} \qquad \hat{\Omega} \vdash \hat{\tau}_2 \rightsquigarrow \tau_2 \Leftarrow \texttt{Type}}{\hat{\Omega} \vdash \texttt{uparr}(\hat{\tau}_1; \hat{\tau}_2) \rightsquigarrow \texttt{parr}(\tau_1; \tau_2) \Rightarrow \texttt{Type}} \tag{5.18g}$$

$$\frac{\hat{\Omega} \vdash \hat{\kappa} \rightsquigarrow \kappa \text{ kind} \qquad \hat{\Omega}, \hat{u} \rightsquigarrow u :: \kappa \vdash \hat{\tau} \rightsquigarrow \tau \Leftarrow \texttt{Type}}{\hat{\Omega} \vdash \texttt{uall}\{\hat{\kappa}\}(\hat{u}.\hat{\tau}) \rightsquigarrow \texttt{all}\{\kappa\}(u.\tau) \Rightarrow \texttt{Type}} \tag{5.18h}$$

$$\frac{\hat{\Omega}, \hat{t} \rightsquigarrow t :: \texttt{Type} \vdash \hat{\tau} \rightsquigarrow \tau \Leftarrow \texttt{Type}}{\hat{\Omega} \vdash \texttt{urec}(\hat{t}.\hat{\tau}) \rightsquigarrow \texttt{rec}(t.\tau) \Rightarrow \texttt{Type}} \tag{5.18i}$$

$$\frac{\{\hat{\Omega} \vdash \hat{\tau}_i \rightsquigarrow \tau_i \Leftarrow \texttt{Type}\}_{1 \le i \le n}}{\hat{\Omega} \vdash \texttt{uprod}[L](\{i \hookrightarrow \hat{\tau}_i\}_{i \in L}) \rightsquigarrow \texttt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L}) \Rightarrow \texttt{Type}} \tag{5.18j}$$

$$\frac{\{\hat{\Omega} \vdash \hat{\tau}_i \rightsquigarrow \tau_i \Leftarrow \texttt{Type}\}_{1 \le i \le n}}{\hat{\Omega} \vdash \texttt{usum}[L](\{i \hookrightarrow \hat{\tau}_i\}_{i \in L}) \rightsquigarrow \texttt{sum}[L](\{i \hookrightarrow \tau_i\}_{i \in L}) \Rightarrow \texttt{Type}} \tag{5.18k}$$

$$\overline{\hat{\Omega}, \hat{X} \rightsquigarrow X : \texttt{sig}\{\kappa\}(u.\tau) \vdash \texttt{ucon}(\hat{X}) \rightsquigarrow \texttt{con}(X) \Rightarrow \kappa} \tag{5.18l}$$

Analytic constructor expansion

$$\frac{\hat{\Omega} \vdash \hat{c} \rightsquigarrow c \Rightarrow \kappa_1 \qquad \Omega \vdash \kappa_1 <:: \kappa_2}{\hat{\Omega} \vdash \hat{c} \rightsquigarrow c \Leftarrow \kappa_2} \tag{5.19a}$$

$$\frac{\hat{\Omega} \vdash \hat{c} \rightsquigarrow c \Leftarrow \texttt{Type}}{\hat{\Omega} \vdash \hat{c} \rightsquigarrow c \Leftarrow \texttt{S}(c)} \tag{5.19b}$$

$$\frac{\hat{\Omega}, \hat{u} \rightsquigarrow u :: \kappa_1 \vdash \hat{c}_2 \rightsquigarrow c_2 \Leftarrow \kappa_2}{\hat{\Omega} \vdash \texttt{uabs}(\hat{u}.\hat{c}_2) \rightsquigarrow \texttt{abs}(u.c_2) \Leftarrow \texttt{darr}(\kappa_1; u.\kappa_2)} \tag{5.19c}$$

$$\frac{\hat{\Omega} \vdash \hat{c}_1 \rightsquigarrow c_1 \Leftarrow \kappa_1 \qquad \hat{\Omega} \vdash \hat{c}_2 \rightsquigarrow c_2 \Leftarrow [c_1/u]\kappa_2}{\hat{\Omega} \vdash \texttt{upair}(\hat{c}_1; \hat{c}_2) \rightsquigarrow \texttt{pair}(c_1; c_2) \Leftarrow \texttt{dprod}(\kappa_1; u.\kappa_2)} \tag{5.19d}$$

Types, Expressions, Rules and Patterns

Type expansion

$$\frac{\hat{\Omega} \vdash \hat{\tau} \rightsquigarrow \tau \Leftarrow \texttt{Type}}{\hat{\Omega} \vdash \hat{\tau} \rightsquigarrow \tau \ \texttt{type}} \tag{5.20}$$

Synthetic typed expression expansion

$$\frac{}{\Omega, \hat{x} \rightsquigarrow x : \tau \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{x} \rightsquigarrow x \Rightarrow \tau} \tag{5.21a}$$

$$\frac{\hat{\Omega} \vdash \hat{\tau} \rightsquigarrow \tau \ \texttt{type} \qquad \hat{\Omega} \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{e} \rightsquigarrow e \Leftarrow \tau}{\hat{\Omega} \vdash_{\hat{\Psi}; \hat{\Phi}} \texttt{uasc}\{\hat{\tau}\}(\hat{e}) \rightsquigarrow e \Rightarrow \tau} \tag{5.21b}$$

$$\frac{\hat{\Omega} \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{e} \rightsquigarrow e \Rightarrow \tau \qquad \Omega, \hat{x} \rightsquigarrow x : \tau \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{e}' \rightsquigarrow e' \Rightarrow \tau'}{\hat{\Omega} \vdash_{\hat{\Psi}; \hat{\Phi}} \texttt{uletval}(\hat{e}; \hat{x}.\hat{e}') \rightsquigarrow \texttt{ap}(\texttt{lam}\{\tau\}(x.e'); e) \Rightarrow \tau'} \tag{5.21c}$$

$$\frac{\hat{\Omega} \vdash \hat{\tau}_1 \rightsquigarrow \tau_1 \ \texttt{type} \qquad \Omega, \hat{x} \rightsquigarrow x : \tau_1 \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{e} \rightsquigarrow e \Rightarrow \tau_2}{\hat{\Omega} \vdash_{\hat{\Psi}; \hat{\Phi}} \texttt{ulam}\{\hat{\tau}_1\}(\hat{x}.\hat{e}) \rightsquigarrow \texttt{lam}\{\tau_1\}(x.e) \Rightarrow \texttt{parr}(\tau_1; \tau_2)} \tag{5.21d}$$

$$\frac{\hat{\Omega} \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{e}_1 \rightsquigarrow e_1 \Rightarrow \texttt{parr}(\tau_2; \tau) \qquad \hat{\Omega} \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{e}_2 \rightsquigarrow e_2 \Leftarrow \tau_2}{\hat{\Omega} \vdash_{\hat{\Psi}; \hat{\Phi}} \texttt{uap}(\hat{e}_1; \hat{e}_2) \rightsquigarrow \texttt{ap}(e_1; e_2) \Rightarrow \tau} \tag{5.21e}$$

$$\frac{\hat{\Omega} \vdash \hat{\kappa} \rightsquigarrow \kappa \ \texttt{kind} \qquad \Omega, \hat{u} \rightsquigarrow u :: \kappa \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{e} \rightsquigarrow e \Rightarrow \tau}{\hat{\Omega} \vdash_{\hat{\Psi}; \hat{\Phi}} \texttt{uclam}\{\hat{\kappa}\}(\hat{u}.\hat{e}) \rightsquigarrow \texttt{clam}\{\kappa\}(u.e) \Rightarrow \texttt{all}\{\kappa\}(u.\tau)} \tag{5.21f}$$

$$\frac{\hat{\Omega} \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{e} \rightsquigarrow e \Rightarrow \texttt{all}\{\kappa\}(u.\tau) \qquad \hat{\Omega} \vdash \hat{c} \rightsquigarrow c \Rightarrow \kappa}{\hat{\Omega} \vdash_{\hat{\Psi}; \hat{\Phi}} \texttt{ucap}\{\hat{c}\}(\hat{e}) \rightsquigarrow \texttt{cap}\{c\}(e) \Rightarrow [c/t]\tau} \tag{5.21g}$$

$$\frac{\hat{\Omega} \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{e} \rightsquigarrow e \Rightarrow \texttt{rec}(t.\tau)}{\hat{\Omega} \vdash_{\hat{\Psi}; \hat{\Phi}} \texttt{uunfold}(\hat{e}) \rightsquigarrow \texttt{unfold}(e) \Rightarrow [\texttt{rec}(t.\tau)/t]\tau} \tag{5.21h}$$

$$\frac{\hat{e} = \texttt{utpl}[L](\{i \hookrightarrow \hat{e}_i\}_{i \in L}) \qquad e = \texttt{tpl}[L](\{i \hookrightarrow e_i\}_{i \in L})}{\{\hat{\Omega} \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{e}_i \rightsquigarrow e_i \Rightarrow \tau_i\}_{i \in L}}{\hat{\Omega} \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{e} \rightsquigarrow e \Rightarrow \texttt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L})} \tag{5.21i}$$

$$\frac{\hat{\Omega} \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{e} \rightsquigarrow e \Rightarrow \texttt{prod}[L, \ell](\{i \hookrightarrow \tau_i\}_{i \in L}; \ell \hookrightarrow \tau)}{\hat{\Omega} \vdash_{\hat{\Psi}; \hat{\Phi}} \texttt{upr}[\ell](\hat{e}) \rightsquigarrow \texttt{pr}[\ell](e) \Rightarrow \tau} \tag{5.21j}$$

$$\frac{n > 0 \qquad \hat{\Omega} \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{e} \rightsquigarrow e \Rightarrow \tau \qquad \{\hat{\Omega} \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{r}_i \rightsquigarrow r_i \Rightarrow \tau \mapsto \tau'\}_{1 \leq i \leq n}}{\hat{\Omega} \vdash_{\hat{\Psi}; \hat{\Phi}} \texttt{umatch}[n](\hat{e}; \{\hat{r}_i\}_{1 \leq i \leq n}) \rightsquigarrow \texttt{match}[n]\{\tau'\}(e; \{r_i\}_{1 \leq i \leq n}) \Rightarrow \tau'} \tag{5.21k}$$

$$\frac{}{\hat{\Omega}, \hat{X} \rightsquigarrow X : \texttt{sig}\{\kappa\}(u.\tau) \vdash_{\hat{\Psi}; \hat{\Phi}} \texttt{uval}(\hat{X}) \rightsquigarrow \texttt{val}(X) \Rightarrow [\texttt{con}(X)/u]\tau} \tag{5.21l}$$

$$\dfrac{\begin{array}{c}\hat{\Omega} \vdash \hat{\rho} \rightsquigarrow \rho \text{ tsmty} \qquad \varnothing \vdash e_{\text{parse}} : \texttt{parr}(\texttt{Body};\texttt{ParseResult}(\texttt{PCEExp})) \\[4pt] \hat{\Omega} \vdash_{\langle \mathcal{A} \uplus \hat{a} \hookrightarrow \texttt{defref}[a];\Psi,a \hookrightarrow \texttt{petsm}(\rho;e_{\text{parse}});\mathcal{I}\rangle;\hat{\Phi}} \hat{e} \rightsquigarrow e \Rightarrow \tau\end{array}}{\hat{\Omega} \vdash_{\langle \mathcal{A};\Psi;\mathcal{I}\rangle;\hat{\Phi}} \texttt{syntax } \hat{a} \texttt{ at } \hat{\rho} \texttt{ for expressions } \{e_{\text{parse}}\} \texttt{ in } \hat{e} \rightsquigarrow e \Rightarrow \tau} \tag{5.21m}$$

$$\dfrac{\hat{\Omega} \vdash^{\textsf{Exp}}_{\langle \mathcal{A};\Psi;\mathcal{I}\rangle} \hat{\epsilon} \rightsquigarrow \epsilon @ \rho \qquad \hat{\Omega} \vdash_{\langle \mathcal{A} \uplus \hat{a} \hookrightarrow \epsilon;\Psi;\mathcal{I}\rangle;\hat{\Phi}} \hat{e} \rightsquigarrow e \Rightarrow \tau}{\hat{\Omega} \vdash_{\langle \mathcal{A};\Psi;\mathcal{I}\rangle;\hat{\Phi}} \texttt{let syntax } \hat{a} = \hat{\epsilon} \texttt{ for expressions in } \hat{e} \rightsquigarrow e \Rightarrow \tau} \tag{5.21n}$$

$$\dfrac{\begin{array}{c}\hat{\Omega} = \langle \mathcal{D};\mathcal{G};\mathcal{M};\Omega_{\text{app}}\rangle \qquad \hat{\Psi} = \langle \mathcal{A};\Psi,a \hookrightarrow \texttt{petsm}(\rho;e_{\text{parse}});\mathcal{I}\rangle \\[4pt] \hat{\Omega} \vdash^{\textsf{Exp}}_{\hat{\Psi}} \hat{\epsilon} \rightsquigarrow \epsilon @ \texttt{type}(\tau_{\text{final}}) \qquad \Omega_{\text{app}} \vdash^{\textsf{Exp}}_{\Psi} \epsilon \Downarrow \epsilon_{\text{normal}} \\[4pt] \texttt{tsmdef}(\epsilon_{\text{normal}}) = a \qquad b \downarrow e_{\text{body}} \qquad e_{\text{parse}}(e_{\text{body}}) \Downarrow \texttt{inj Success} \cdot e_{\text{pcand}} \\[4pt] e_{\text{pcand}} \uparrow_{\texttt{PCEExp}} \check{e} \\[4pt] \Omega_{\text{app}} \vdash^{\textsf{Exp}}_{\Psi,a \hookrightarrow \texttt{petsm}(\rho;e_{\text{parse}})} \check{e} \overset{\looparrowright}{\rightarrow}_{\epsilon_{\text{normal}}} \grave{e} ? \texttt{type}(\tau_{\text{cand}}) \dashv \omega : \Omega_{\text{params}} \\[4pt] \Omega_{\text{params}} \vdash^{\hat{\Omega};\hat{\Psi};\hat{\Phi};b} \grave{e} \rightsquigarrow e \Leftarrow \tau_{\text{cand}}\end{array}}{\hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \ /b/ \rightsquigarrow [\omega]e \Rightarrow [\omega]\tau_{\text{cand}}} \tag{5.21o}$$

TODO: peTSM implicit designation

$$\dfrac{\hat{\Delta}\,\hat{\Gamma} \vdash_{\langle \mathcal{A} \uplus \hat{a} \rightsquigarrow a;\Psi,a \hookrightarrow \texttt{uetsm}(\tau;e_{\text{parse}});\mathcal{I} \uplus \tau \hookrightarrow a\rangle;\hat{\Phi}} \hat{e} \rightsquigarrow e \Rightarrow \tau'}{\hat{\Delta}\,\hat{\Gamma} \vdash_{\langle \mathcal{A} \uplus \hat{a} \rightsquigarrow a;\Psi,a \hookrightarrow \texttt{uetsm}(\tau;e_{\text{parse}});\mathcal{I}\rangle;\hat{\Phi}} \texttt{implicit syntax } \hat{a} \texttt{ for expressions in } \hat{e} \rightsquigarrow e \Rightarrow \tau'} \tag{5.21p}$$

$$\dfrac{\begin{array}{c}\hat{\Omega} \vdash \hat{\rho} \rightsquigarrow \rho \text{ tsmty} \qquad \varnothing \vdash e_{\text{parse}} : \texttt{parr}(\texttt{Body};\texttt{ParseResult}(\texttt{CEPat})) \\[4pt] \hat{\Omega} \vdash_{\hat{\Psi};\langle \mathcal{A} \uplus \hat{a} \hookrightarrow \texttt{defref}[a];\Phi,a \hookrightarrow \texttt{pptsm}(\rho;e_{\text{parse}});\mathcal{I}\rangle} \hat{e} \rightsquigarrow e \Rightarrow \tau\end{array}}{\hat{\Omega} \vdash_{\hat{\Psi};\langle \mathcal{A};\Phi;\mathcal{I}\rangle} \texttt{syntax } \hat{a} \texttt{ at } \hat{\rho} \texttt{ for patterns } \{e_{\text{parse}}\} \texttt{ in } \hat{e} \rightsquigarrow e \Rightarrow \tau} \tag{5.21q}$$

$$\dfrac{\hat{\Omega} \vdash^{\textsf{Pat}}_{\langle \mathcal{A};\Phi;\mathcal{I}\rangle} \hat{\epsilon} \rightsquigarrow \epsilon @ \rho \qquad \hat{\Omega} \vdash_{\hat{\Psi};\langle \mathcal{A} \uplus \hat{a} \hookrightarrow \epsilon;\Phi;\mathcal{I}\rangle} \hat{e} \rightsquigarrow e \Rightarrow \tau}{\hat{\Omega} \vdash_{\hat{\Psi};\langle \mathcal{A};\Phi;\mathcal{I}\rangle} \texttt{let syntax } \hat{a} = \hat{\epsilon} \texttt{ for patterns in } \hat{e} \rightsquigarrow e \Rightarrow \tau} \tag{5.21r}$$

TODO: ppTSM implicit designation

$$\dfrac{\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\langle \mathcal{A} \uplus \hat{a} \rightsquigarrow a;\Phi,a \hookrightarrow \texttt{uptsm}(\tau;e_{\text{parse}});\mathcal{I} \uplus \tau \hookrightarrow a\rangle} \hat{e} \rightsquigarrow e \Rightarrow \tau'}{\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\langle \mathcal{A} \uplus \hat{a} \rightsquigarrow a;\Phi,a \hookrightarrow \texttt{uetsm}(\tau;e_{\text{parse}});\mathcal{I}\rangle} \texttt{implicit syntax } \hat{a} \texttt{ for patterns in } \hat{e} \rightsquigarrow e \Rightarrow \tau'} \tag{5.21s}$$

Analytic typed expression expansion

$$\dfrac{\hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \rightsquigarrow e \Rightarrow \tau \qquad \Omega \vdash \tau <: \tau'}{\hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \rightsquigarrow e \Leftarrow \tau'} \tag{5.22a}$$

$$\frac{\hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \leadsto e \Rightarrow \tau \qquad \hat{\Omega}, \hat{x} \leadsto x : \tau \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e}' \leadsto e' \Leftarrow \tau'}{\hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \mathtt{uletval}(\hat{e}; \hat{x}.\hat{e}') \leadsto \mathtt{ap}(\mathtt{lam}\{\tau\}(x.e'); e) \Leftarrow \tau'} \tag{5.22b}$$

$$\frac{\hat{\Omega}, \hat{x} \leadsto x : \tau_1 \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \leadsto e \Leftarrow \tau_2}{\hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \mathtt{uanalam}(\hat{x}.\hat{e}) \leadsto \mathtt{lam}\{\tau_1\}(x.e) \Leftarrow \mathtt{parr}(\tau_1; \tau_2)} \tag{5.22c}$$

$$\frac{\hat{\Delta}, \hat{t} \leadsto t \; \mathsf{type} \; \hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \leadsto e \Leftarrow \tau}{\hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \mathtt{utlam}(\hat{t}.\hat{e}) \leadsto \mathtt{tlam}(t.e) \Leftarrow \mathtt{all}(t.\tau)} \tag{5.22d}$$

$$\frac{\hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \leadsto e \Leftarrow [\mathtt{rec}(t.\tau)/t]\tau}{\hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \mathtt{ufold}(\hat{e}) \leadsto \mathtt{fold}\{t.\tau\}(e) \Leftarrow \mathtt{rec}(t.\tau)} \tag{5.22e}$$

$$\frac{\begin{array}{c}\hat{e} = \mathtt{utpl}[L](\{i \hookrightarrow \hat{e}_i\}_{i \in L}) \qquad e = \mathtt{tpl}[L](\{i \hookrightarrow e_i\}_{i \in L}) \\ \{\hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e}_i \leadsto e_i \Leftarrow \tau_i\}_{i \in L}\end{array}}{\hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \leadsto e \Leftarrow \mathtt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L})} \tag{5.22f}$$

$$\frac{\begin{array}{c}\hat{e} = \mathtt{uin}[\ell](\hat{e}') \qquad e = \mathtt{in}[L, \ell][\ell]\{\{i \hookrightarrow \tau_i\}_{i \in L}; \ell \hookrightarrow \tau\}(e') \\ \hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e}' \leadsto e' \Leftarrow \tau\end{array}}{\hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \leadsto e \Leftarrow \mathtt{sum}[L, \ell](\{i \hookrightarrow \tau_i\}_{i \in L}; \ell \hookrightarrow \tau)} \tag{5.22g}$$

$$\frac{\hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \leadsto e \Rightarrow \tau \qquad \{\hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{r}_i \leadsto r_i \Leftarrow \tau \mapsto \tau'\}_{1 \leq i \leq n}}{\hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \mathtt{umatch}[n](\hat{e}; \{\hat{r}_i\}_{1 \leq i \leq n}) \leadsto \mathtt{match}[n]\{\tau'\}(e; \{r_i\}_{1 \leq i \leq n}) \Leftarrow \tau'} \tag{5.22h}$$

$$\frac{\begin{array}{c}\hat{\Omega} \vdash \hat{\rho} \leadsto \rho \; \mathsf{tsmty} \qquad \varnothing \vdash e_{\mathrm{parse}} : \mathtt{parr}(\mathsf{Body}; \mathsf{ParseResult}(\mathsf{PCEExp})) \\ \hat{\Omega} \vdash_{\langle \mathcal{A} \uplus \hat{a} \hookrightarrow \mathtt{defref}[a]; \Psi, a \hookrightarrow \mathtt{petsm}(\rho; e_{\mathrm{parse}}); \mathcal{I}\rangle; \hat{\Phi}} \hat{e} \leadsto e \Leftarrow \tau\end{array}}{\hat{\Omega} \vdash_{\langle \mathcal{A}; \Psi; \mathcal{I}\rangle; \hat{\Phi}} \mathtt{syntax} \; \hat{a} \; \mathtt{at} \; \hat{\rho} \; \mathtt{for} \; \mathtt{expressions} \; \{e_{\mathrm{parse}}\} \; \mathtt{in} \; \hat{e} \leadsto e \Leftarrow \tau} \tag{5.22i}$$

$$\frac{\hat{\Omega} \vdash^{\mathsf{Exp}}_{\langle \mathcal{A}; \Psi; \mathcal{I}\rangle} \hat{e} \leadsto \epsilon @ \rho \qquad \hat{\Omega} \vdash_{\langle \mathcal{A} \uplus \hat{a} \hookrightarrow \epsilon; \Psi; \mathcal{I}\rangle; \hat{\Phi}} \hat{e} \leadsto e \Leftarrow \tau}{\hat{\Omega} \vdash_{\langle \mathcal{A}; \Psi; \mathcal{I}\rangle; \hat{\Phi}} \mathtt{let} \; \mathtt{syntax} \; \hat{a} = \hat{e} \; \mathtt{for} \; \mathtt{expressions} \; \mathtt{in} \; \hat{e} \leadsto e \Leftarrow \tau} \tag{5.22j}$$

TODO: peTSM implicit designation

$$\frac{\hat{\Delta} \; \hat{\Gamma} \vdash_{\langle \mathcal{A} \uplus \hat{a} \leadsto a; \Psi, a \hookrightarrow \mathtt{uetsm}(\tau; e_{\mathrm{parse}}); \mathcal{I} \uplus \tau \hookrightarrow a\rangle; \hat{\Phi}} \hat{e} \leadsto e \Leftarrow \tau'}{\hat{\Delta} \; \hat{\Gamma} \vdash_{\langle \mathcal{A} \uplus \hat{a} \leadsto a; \Psi, a \hookrightarrow \mathtt{uetsm}(\tau; e_{\mathrm{parse}}); \mathcal{I}\rangle; \hat{\Phi}} \mathtt{implicit} \; \mathtt{syntax} \; \hat{a} \; \mathtt{for} \; \mathtt{expressions} \; \mathtt{in} \; \hat{e} \leadsto e \Leftarrow \tau'} \tag{5.22k}$$

TODO: peTSM implicit application

$$\frac{\begin{array}{c}b \downarrow e_{\mathrm{body}} \qquad e_{\mathrm{parse}}(e_{\mathrm{body}}) \Downarrow \mathtt{inj} \; \mathsf{Success} \cdot e_{\mathrm{cand}} \qquad e_{\mathrm{cand}} \uparrow_{\mathsf{CEExp}} \grave{e} \\ \varnothing \; \varnothing \vdash^{\hat{\Delta}; \hat{\Gamma}; \langle \mathcal{A}; \Psi, a \hookrightarrow \mathtt{uetsm}(\tau; e_{\mathrm{parse}}); \mathcal{I} \uplus \tau \hookrightarrow a\rangle; \hat{\Phi}; b} \grave{e} \leadsto e \Leftarrow \tau\end{array}}{\hat{\Delta} \; \hat{\Gamma} \vdash_{\langle \mathcal{A}; \Psi, a \hookrightarrow \mathtt{uetsm}(\tau; e_{\mathrm{parse}}); \mathcal{I} \uplus \tau \hookrightarrow a\rangle; \hat{\Phi}} \mathtt{uelit}[b] \leadsto e \Leftarrow \tau} \tag{5.22l}$$

$$\dfrac{\begin{array}{c}\hat{\Omega} \vdash \hat{\rho} \rightsquigarrow \rho \text{ tsmty} \qquad \varnothing \vdash e_{\text{parse}} : \mathtt{parr}(\mathtt{Body}; \mathtt{ParseResult}(\mathtt{CEPat})) \\ \hat{\Omega} \vdash_{\hat{\Psi};\langle \mathcal{A} \uplus \hat{a} \hookrightarrow \mathtt{defref}[a];\Phi,a \hookrightarrow \mathtt{pptsm}(\rho;e_{\text{parse}});\mathcal{I}\rangle} \hat{e} \rightsquigarrow e \Leftarrow \tau \end{array}}{\hat{\Omega} \vdash_{\hat{\Psi};\langle \mathcal{A};\Phi;\mathcal{I}\rangle} \mathtt{syntax}\ \hat{a}\ \mathtt{at}\ \hat{\rho}\ \mathtt{for\ patterns}\ \{e_{\text{parse}}\}\ \mathtt{in}\ \hat{e} \rightsquigarrow e \Leftarrow \tau} \tag{5.22m}$$

$$\dfrac{\hat{\Omega} \vdash^{\mathsf{Pat}}_{\langle \mathcal{A};\Phi;\mathcal{I}\rangle} \hat{e} \rightsquigarrow \epsilon @ \rho \qquad \hat{\Omega} \vdash_{\hat{\Psi};\langle \mathcal{A} \uplus \hat{a} \hookrightarrow \epsilon;\Phi;\mathcal{I}\rangle} \hat{e} \rightsquigarrow e \Leftarrow \tau}{\hat{\Omega} \vdash_{\hat{\Psi};\langle \mathcal{A};\Phi;\mathcal{I}\rangle} \mathtt{let\ syntax}\ \hat{a} = \hat{e}\ \mathtt{for\ patterns\ in}\ \hat{e} \rightsquigarrow e \Leftarrow \tau} \tag{5.22n}$$

TODO: ppTSM implicit designation

$$\dfrac{\hat{\Delta}\ \hat{\Gamma} \vdash_{\hat{\Psi};\langle \mathcal{A} \uplus \hat{a} \rightsquigarrow a;\Phi,a \hookrightarrow \mathtt{uptsm}(\tau;e_{\text{parse}});\mathcal{I} \uplus \tau \hookrightarrow a\rangle} \hat{e} \rightsquigarrow e \Leftarrow \tau'}{\hat{\Delta}\ \hat{\Gamma} \vdash_{\hat{\Psi};\langle \mathcal{A} \uplus \hat{a} \rightsquigarrow a;\Phi,a \hookrightarrow \mathtt{uetsm}(\tau;e_{\text{parse}});\mathcal{I}\rangle} \mathtt{implicit\ syntax}\ \hat{a}\ \mathtt{for\ patterns\ in}\ \hat{e} \rightsquigarrow e \Leftarrow \tau'} \tag{5.22o}$$

Synthetic rule expansion

$$\dfrac{\begin{array}{c}\hat{\Omega} = \langle \mathcal{D};\mathcal{G};\mathcal{M};\Omega \rangle \\ \hat{\Omega} \vdash_{\hat{\Phi}} \hat{p} \rightsquigarrow p : \tau \dashv\mathbin{\parallel} \langle \varnothing;\mathcal{G}';\varnothing;\Omega' \rangle \qquad \langle \mathcal{D};\mathcal{G} \uplus \mathcal{G}';\mathcal{M};\Omega \cup \Omega' \rangle \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \rightsquigarrow e \Rightarrow \tau' \end{array}}{\hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \mathtt{urule}(\hat{p}.\hat{e}) \rightsquigarrow \mathtt{rule}(p.e) \Rightarrow \tau \mapsto \tau'} \tag{5.23}$$

Analytic rule expansion

$$\dfrac{\begin{array}{c}\hat{\Omega} = \langle \mathcal{D};\mathcal{G};\mathcal{M};\Omega \rangle \\ \hat{\Omega} \vdash_{\hat{\Phi}} \hat{p} \rightsquigarrow p : \tau \dashv\mathbin{\parallel} \langle \varnothing;\mathcal{G}';\varnothing;\Omega' \rangle \qquad \langle \mathcal{D};\mathcal{G} \uplus \mathcal{G}';\mathcal{M};\Omega \cup \Omega' \rangle \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \rightsquigarrow e \Leftarrow \tau' \end{array}}{\hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \mathtt{urule}(\hat{p}.\hat{e}) \rightsquigarrow \mathtt{rule}(p.e) \Leftarrow \tau \mapsto \tau'} \tag{5.24}$$

Typed pattern expansion

$$\dfrac{\begin{array}{c}\hat{\Omega} = \langle \mathcal{D};\mathcal{G};\mathcal{M};\Omega \rangle \\ \hat{\Omega} \vdash_{\hat{\Phi}} \hat{p} \rightsquigarrow p : \tau \dashv\mathbin{\parallel} \hat{\Omega}' \qquad \Omega \vdash \tau <: \tau' \end{array}}{\hat{\Omega} \vdash_{\hat{\Phi}} \hat{p} \rightsquigarrow p : \tau' \dashv\mathbin{\parallel} \hat{\Omega}'} \tag{5.25a}$$

$$\dfrac{}{\hat{\Omega} \vdash_{\hat{\Phi}} \hat{x} \rightsquigarrow x : \tau \dashv\mathbin{\parallel} \langle \varnothing;\hat{x} \rightsquigarrow x;\varnothing;x : \tau \rangle} \tag{5.25b}$$

$$\dfrac{}{\hat{\Omega} \vdash_{\hat{\Phi}} \mathtt{uwildp} \rightsquigarrow \mathtt{wildp} : \tau \dashv\mathbin{\parallel} \langle \varnothing;\varnothing;\varnothing;\varnothing \rangle} \tag{5.25c}$$

$$\dfrac{\hat{\Omega} \vdash_{\hat{\Phi}} \hat{p} \rightsquigarrow p : [\mathtt{rec}(t.\tau)/t]\tau \dashv\mathbin{\parallel} \hat{\Omega}'}{\hat{\Omega} \vdash_{\hat{\Phi}} \mathtt{ufoldp}(\hat{p}) \rightsquigarrow \mathtt{foldp}(p) : \mathtt{rec}(t.\tau) \dashv\mathbin{\parallel} \hat{\Omega}'} \tag{5.25d}$$

$$\dfrac{\begin{array}{c}\hat{p} = \mathtt{utplp}[L](\{i \hookrightarrow \hat{p}_i\}_{i \in L}) \qquad p = \mathtt{tplp}[L](\{i \hookrightarrow p_i\}_{i \in L}) \\ \{\hat{\Omega} \vdash_{\hat{\Phi}} \hat{p}_i \rightsquigarrow p_i : \tau_i \dashv\mathbin{\parallel} \hat{\Omega}_i\}_{i \in L} \end{array}}{\hat{\Omega} \vdash_{\hat{\Phi}} \hat{p} \rightsquigarrow p : \mathtt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L}) \dashv\mathbin{\parallel} \cup_{i \in L} \hat{\Omega}_i} \tag{5.25e}$$

$$\frac{\hat{\Omega} \vdash_{\hat{\Phi}} \hat{p} \rightsquigarrow p : \tau \dashv\!\!\mid \hat{\Omega}'}{\hat{\Omega} \vdash_{\hat{\Phi}} \texttt{uinp}[\ell](\hat{p}) \rightsquigarrow \texttt{inp}[\ell](p) : \texttt{sum}[L, \ell](\{i \hookrightarrow \tau_i\}_{i \in L}; \ell \hookrightarrow \tau) \dashv\!\!\mid \hat{\Omega}'} \quad (5.25\text{f})$$

$$\frac{\begin{array}{c} \hat{\Omega} = \langle \mathcal{D}; \mathcal{G}; \mathcal{M}; \Omega_{\text{app}} \rangle \qquad \hat{\Phi} = \langle \mathcal{A}; \Phi, a \hookrightarrow \texttt{pptsm}(\rho; e_{\text{parse}}); \mathcal{I} \rangle \\ \hat{\Omega} \vdash_{\hat{\Phi}}^{\text{Pat}} \hat{\epsilon} \rightsquigarrow \epsilon @ \texttt{type}(\tau_{\text{final}}) \qquad \texttt{tsmdef}(\epsilon) = a \\ b \downarrow e_{\text{body}} \qquad e_{\text{parse}}(e_{\text{body}}) \Downarrow \texttt{inj Success} \cdot e_{\text{cand}} \qquad e_{\text{cand}} \uparrow_{\text{CEPat}} \hat{p} \\ \Omega_{\text{app}} \vdash_{\Phi, a \hookrightarrow \texttt{pptsm}(\rho; e_{\text{parse}})}^{\text{Pat}} \cdot \looparrowright_\epsilon \cdot ? \texttt{type}(\tau_{\text{cand}}) \dashv \omega : \Omega_{\text{params}} \\ \Omega_{\text{param}} \vdash^{\hat{\Omega}; \hat{\Phi}; b} \hat{p} \rightsquigarrow p : \tau_{\text{cand}} \dashv\!\!\mid \hat{\Omega}' \end{array}}{\hat{\Omega} \vdash_{\hat{\Phi}} \hat{\epsilon} \; /b/ \rightsquigarrow p : [\omega] \tau_{\text{cand}} \dashv\!\!\mid \hat{\Omega}'} \quad (5.25\text{g})$$

TODO: ppTSM implicit application

$$\frac{\begin{array}{c} b \downarrow e_{\text{body}} \qquad e_{\text{parse}}(e_{\text{body}}) \Downarrow \texttt{inj Success} \cdot e_{\text{cand}} \qquad e_{\text{cand}} \uparrow_{\text{CEPat}} \hat{p} \\ \Omega_{\text{param}} \vdash^{\Delta; \langle \mathcal{A}; \Phi, a \hookrightarrow \texttt{uptsm}(\tau; e_{\text{parse}}); \mathcal{I}, \tau \hookrightarrow a \rangle; b} \hat{p} \rightsquigarrow p : \tau \dashv\!\!\mid \hat{\Omega} \end{array}}{\Delta \vdash_{\langle \mathcal{A}; \Phi, a \hookrightarrow \texttt{uptsm}(\tau; e_{\text{parse}}); \mathcal{I}, \tau \hookrightarrow a \rangle} \; /b/ \rightsquigarrow p : \tau \dashv\!\!\mid \hat{\Omega}} \quad (5.25\text{h})$$

Unexpanded Signatures and Module Expressions

Signature expansion

$$\frac{\hat{\Omega} \vdash \hat{\kappa} \rightsquigarrow \kappa \; \mathsf{kind} \qquad \hat{\Omega}, \hat{u} \rightsquigarrow u :: \kappa \vdash \hat{\tau} \rightsquigarrow \tau \; \mathsf{type}}{\hat{\Omega} \vdash \texttt{usig}\{\hat{\kappa}\}(\hat{u}.\hat{\tau}) \rightsquigarrow \texttt{sig}\{\kappa\}(u.\tau) \; \mathsf{sig}} \quad (5.26)$$

Synthetic module expression expansion

$$\frac{}{\hat{\Omega}, \hat{X} \rightsquigarrow X : \sigma \vdash_{\hat{\Phi}; \hat{\Psi}} \hat{X} \rightsquigarrow X \Rightarrow \sigma} \quad (5.27\text{a})$$

$$\frac{\hat{\Omega} \vdash \hat{\sigma} \rightsquigarrow \sigma \; \mathsf{sig} \qquad \hat{\Omega} \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{M} \rightsquigarrow M \Leftarrow \sigma}{\hat{\Omega} \vdash_{\hat{\Psi}_{\text{E}}; \hat{\Psi}_{\text{P}}} \texttt{useal}\{\hat{\sigma}\}(\hat{M}) \rightsquigarrow \texttt{seal}\{\sigma\}(M) \Rightarrow \sigma} \quad (5.27\text{b})$$

$$\frac{\begin{array}{c} \hat{\Omega} \vdash_{\hat{\Psi}_{\text{E}}; \hat{\Psi}_{\text{P}}} \hat{M} \rightsquigarrow M \Rightarrow \sigma \qquad \hat{\Omega} \vdash \hat{\sigma}' \rightsquigarrow \sigma' \; \mathsf{sig} \\ \hat{\Omega}, \hat{X} \rightsquigarrow X : \sigma \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{M}' \rightsquigarrow M' \Leftarrow \sigma' \end{array}}{\hat{\Omega} \vdash_{\hat{\Psi}_{\text{E}}; \hat{\Psi}_{\text{P}}} \texttt{umlet}\{\hat{\sigma}'\}(\hat{M}; \hat{X}.\hat{M}') \rightsquigarrow \texttt{mlet}\{\sigma'\}(M; X.M') \Rightarrow \sigma'} \quad (5.27\text{c})$$

$$\frac{\begin{array}{c} \hat{\Omega} \vdash \hat{\rho} \rightsquigarrow \rho \; \mathsf{tsmty} \qquad \varnothing \vdash e_{\text{parse}} : \texttt{parr}(\texttt{Body}; \texttt{ParseResult}(\texttt{PCEExp})) \\ \hat{\Omega} \vdash_{\langle \mathcal{A} \uplus \hat{a} \hookrightarrow \texttt{defref}[a]; \Psi, a \hookrightarrow \texttt{petsm}(\rho; e_{\text{parse}}); \mathcal{I} \rangle; \hat{\Phi}} \hat{M} \rightsquigarrow M \Rightarrow \sigma \end{array}}{\hat{\Omega} \vdash_{\langle \mathcal{A}; \Psi; \mathcal{I} \rangle; \hat{\Phi}} \texttt{syntax} \; \hat{a} \; \texttt{at} \; \hat{\rho} \; \texttt{for expressions} \; \{e_{\text{parse}}\} \; \texttt{in} \; \hat{M} \rightsquigarrow M \Rightarrow \sigma} \quad (5.27\text{d})$$

$$\frac{\hat{\Omega} \vdash_{\langle \mathcal{A}; \Psi; \mathcal{I} \rangle}^{\text{Exp}} \hat{\epsilon} \rightsquigarrow \epsilon @ \rho \qquad \hat{\Omega} \vdash_{\langle \mathcal{A} \uplus \hat{a} \hookrightarrow \epsilon; \Psi; \mathcal{I} \rangle; \hat{\Phi}} \hat{M} \rightsquigarrow M \Rightarrow \sigma}{\hat{\Omega} \vdash_{\langle \mathcal{A}; \Psi; \mathcal{I} \rangle; \hat{\Phi}} \texttt{let syntax} \; \hat{a} = \hat{\epsilon} \; \texttt{for expressions in} \; \hat{M} \rightsquigarrow M \Rightarrow \sigma} \quad (5.27\text{e})$$

$$\begin{array}{c} \vdots \\ \cdots \end{array} \tag{5.27f}$$

$$\dfrac{\hat{\Omega} \vdash \hat{\rho} \rightsquigarrow \rho \; \mathsf{tsmty} \qquad \varnothing \vdash e_{\mathrm{parse}} : \mathtt{parr}(\mathtt{Body};\mathtt{ParseResult}(\mathtt{CEPat}))}{\hat{\Omega} \vdash_{\hat{\Psi};\langle \mathcal{A} \uplus \hat{a} \hookrightarrow \mathtt{defref}[a];\Phi,a \hookrightarrow \mathtt{pptsm}(\rho;e_{\mathrm{parse}});\mathcal{I} \rangle} \hat{M} \rightsquigarrow M \Rightarrow \sigma}{\hat{\Omega} \vdash_{\hat{\Psi};\langle \mathcal{A};\Phi;\mathcal{I} \rangle} \mathtt{syntax}\; \hat{a} \;\mathtt{at}\; \hat{\rho} \;\mathtt{for}\; \mathtt{patterns}\; \{e_{\mathrm{parse}}\} \;\mathtt{in}\; \hat{M} \rightsquigarrow M \Rightarrow \sigma} \tag{5.27g}$$

$$\dfrac{\hat{\Omega} \vdash^{\mathsf{Pat}}_{\langle \mathcal{A};\Phi;\mathcal{I} \rangle} \hat{\epsilon} \rightsquigarrow \epsilon \,@\, \rho \qquad \hat{\Omega} \vdash_{\hat{\Psi};\langle \mathcal{A} \uplus \hat{a} \hookrightarrow \epsilon;\Phi;\mathcal{I} \rangle} \hat{M} \rightsquigarrow M \Rightarrow \sigma}{\hat{\Omega} \vdash_{\hat{\Psi};\langle \mathcal{A};\Phi;\mathcal{I} \rangle} \mathtt{let}\;\mathtt{syntax}\; \hat{a} = \hat{\epsilon} \;\mathtt{for}\; \mathtt{patterns}\; \mathtt{in}\; \hat{M} \rightsquigarrow M \Rightarrow \sigma} \tag{5.27h}$$

$$\begin{array}{c} \vdots \\ \cdots \end{array} \tag{5.27i}$$

Analytic module expression expansion

$$\dfrac{\hat{\Omega} \vdash_{\hat{\Psi}_{\mathrm{E}};\hat{\Psi}_{\mathrm{P}}} \hat{M} \rightsquigarrow M \Rightarrow \sigma \qquad \hat{\Omega} \vdash \sigma <: \sigma'}{\hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{M} \rightsquigarrow M \Leftarrow \sigma'} \tag{5.28a}$$

$$\dfrac{\hat{\Omega} \vdash \hat{c} \rightsquigarrow c \Leftarrow \kappa \qquad \hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \rightsquigarrow e \Leftarrow [c/u]\tau}{\hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \mathtt{ustruct}(\hat{c};\hat{e}) \rightsquigarrow \mathtt{struct}(c;e) \Leftarrow \mathtt{sig}\{\kappa\}(u.\tau)} \tag{5.28b}$$

$$\dfrac{\hat{\Omega} \vdash \hat{\rho} \rightsquigarrow \rho \; \mathsf{tsmty} \qquad \varnothing \vdash e_{\mathrm{parse}} : \mathtt{parr}(\mathtt{Body};\mathtt{ParseResult}(\mathtt{PCEExp}))}{\hat{\Omega} \vdash_{\langle \mathcal{A} \uplus \hat{a} \hookrightarrow \mathtt{defref}[a];\Psi,a \hookrightarrow \mathtt{petsm}(\rho;e_{\mathrm{parse}});\mathcal{I} \rangle;\hat{\Phi}} \hat{M} \rightsquigarrow M \Leftarrow \sigma}{\hat{\Omega} \vdash_{\langle \mathcal{A};\Psi;\mathcal{I} \rangle;\hat{\Phi}} \mathtt{syntax}\; \hat{a} \;\mathtt{at}\; \hat{\rho} \;\mathtt{for}\; \mathtt{expressions}\; \{e_{\mathrm{parse}}\} \;\mathtt{in}\; \hat{M} \rightsquigarrow M \Leftarrow \sigma} \tag{5.28c}$$

$$\dfrac{\hat{\Omega} \vdash^{\mathsf{Exp}}_{\langle \mathcal{A};\Psi;\mathcal{I} \rangle} \hat{\epsilon} \rightsquigarrow \epsilon \,@\, \rho \qquad \hat{\Omega} \vdash_{\langle \mathcal{A} \uplus \hat{a} \hookrightarrow \epsilon;\Psi;\mathcal{I} \rangle;\hat{\Phi}} \hat{M} \rightsquigarrow M \Leftarrow \sigma}{\hat{\Omega} \vdash_{\langle \mathcal{A};\Psi;\mathcal{I} \rangle;\hat{\Phi}} \mathtt{let}\;\mathtt{syntax}\; \hat{a} = \hat{\epsilon} \;\mathtt{for}\; \mathtt{expressions}\; \mathtt{in}\; \hat{M} \rightsquigarrow M \Leftarrow \sigma} \tag{5.28d}$$

$$\begin{array}{c} \vdots \\ \cdots \end{array} \tag{5.28e}$$

$$\dfrac{\hat{\Omega} \vdash \hat{\rho} \rightsquigarrow \rho \; \mathsf{tsmty} \qquad \varnothing \vdash e_{\mathrm{parse}} : \mathtt{parr}(\mathtt{Body};\mathtt{ParseResult}(\mathtt{CEPat}))}{\hat{\Omega} \vdash_{\hat{\Psi};\langle \mathcal{A} \uplus \hat{a} \hookrightarrow \mathtt{defref}[a];\Phi,a \hookrightarrow \mathtt{pptsm}(\rho;e_{\mathrm{parse}});\mathcal{I} \rangle} \hat{M} \rightsquigarrow M \Leftarrow \sigma}{\hat{\Omega} \vdash_{\hat{\Psi};\langle \mathcal{A};\Phi;\mathcal{I} \rangle} \mathtt{syntax}\; \hat{a} \;\mathtt{at}\; \hat{\rho} \;\mathtt{for}\; \mathtt{patterns}\; \{e_{\mathrm{parse}}\} \;\mathtt{in}\; \hat{M} \rightsquigarrow M \Leftarrow \sigma} \tag{5.28f}$$

$$\dfrac{\hat{\Omega} \vdash^{\mathsf{Pat}}_{\langle \mathcal{A};\Phi;\mathcal{I} \rangle} \hat{\epsilon} \rightsquigarrow \epsilon \,@\, \rho \qquad \hat{\Omega} \vdash_{\hat{\Psi};\langle \mathcal{A} \uplus \hat{a} \hookrightarrow \epsilon;\Phi;\mathcal{I} \rangle} \hat{M} \rightsquigarrow M \Leftarrow \sigma}{\hat{\Omega} \vdash_{\hat{\Psi};\langle \mathcal{A};\Phi;\mathcal{I} \rangle} \mathtt{let}\;\mathtt{syntax}\; \hat{a} = \hat{\epsilon} \;\mathtt{for}\; \mathtt{patterns}\; \mathtt{in}\; \hat{M} \rightsquigarrow M \Leftarrow \sigma} \tag{5.28g}$$

$$\begin{array}{c} \vdots \\ \cdots \end{array} \tag{5.28h}$$

TSM Types and Expressions

TSM Expression Typing

Judgement Form	Description
$\Omega \vdash \rho$ tsmty	ρ is a well-formed TSM type
$\Omega \vdash^{\mathsf{Exp}}_{\Psi} \epsilon \,@\, \rho$	peTSM expression ϵ has TSM type ρ
$\Omega \vdash^{\mathsf{Pat}}_{\Phi} \epsilon \,@\, \rho$	ppTSM expression ϵ has TSM type ρ

peTSM Expression Evaluation

Judgement Form	Description
$\Omega \vdash^{\mathsf{Exp}}_{\Psi} \epsilon$ normal	peTSM expression ϵ is in normal form
$\Omega \vdash^{\mathsf{Exp}}_{\Psi} \epsilon \mapsto \epsilon'$	peTSM expression ϵ transitions to ϵ'

+ auxiliary judgements for multi-step transitions and evaluation
unexpanded TSM types and expressions

Judgement Form	Description
$\hat{\Omega} \vdash \hat{\rho} \rightsquigarrow \rho$ tsmty	$\hat{\rho}$ has expansion ρ
$\hat{\Omega} \vdash^{\mathsf{Exp}}_{\hat{\Psi}} \hat{\epsilon} \rightsquigarrow \epsilon \,@\, \rho$	unexpanded peTSM expression $\hat{\epsilon}$ has expansion ϵ and type ρ
$\hat{\Omega} \vdash^{\mathsf{Pat}}_{\hat{\Phi}} \hat{\epsilon} \rightsquigarrow \epsilon \,@\, \rho$	unexpanded ppTSM expression $\hat{\epsilon}$ has expansion ϵ and type ρ

TSM type formation

$$\frac{\Omega \vdash \tau \text{ type}}{\Omega \vdash \mathtt{type}(\tau) \text{ tsmty}} \tag{5.29a}$$

$$\frac{\Omega, t :: \mathsf{Type} \vdash \rho \text{ tsmty}}{\Omega \vdash \mathtt{alltypes}(t.\rho) \text{ tsmty}} \tag{5.29b}$$

$$\frac{\Omega \vdash \sigma \text{ sig} \qquad \Omega, X : \sigma \vdash \rho \text{ tsmty}}{\Omega \vdash \mathtt{allmods}\{\sigma\}(X.\rho) \text{ tsmty}} \tag{5.29c}$$

Unexpanded TSM type expansion

$$\frac{\hat{\Omega} \vdash \hat{\tau} \rightsquigarrow \tau \text{ type}}{\hat{\Omega} \vdash \mathtt{utype}(\hat{\tau}) \rightsquigarrow \mathtt{type}(\tau) \text{ tsmty}} \tag{5.30a}$$

$$\frac{\hat{\Omega}, \hat{t} \rightsquigarrow t :: \mathsf{Type} \vdash \hat{\rho} \rightsquigarrow \rho \text{ tsmty}}{\hat{\Omega} \vdash \mathtt{ualltypes}(\hat{t}.\hat{\rho}) \rightsquigarrow \mathtt{alltypes}(t.\rho) \text{ tsmty}} \tag{5.30b}$$

$$\frac{\hat{\Omega} \vdash \hat{\sigma} \rightsquigarrow \sigma \text{ sig} \qquad \hat{\Omega}, \hat{X} \rightsquigarrow X : \sigma \vdash \hat{\rho} \rightsquigarrow \rho \text{ tsmty}}{\hat{\Omega} \vdash \mathtt{uallmods}\{\hat{\sigma}\}(\hat{X}.\hat{\rho}) \rightsquigarrow \mathtt{allmods}\{\sigma\}(X.\rho) \text{ tsmty}} \tag{5.30c}$$

peTSM Expression Typing

$$\frac{}{\Omega \vdash^{\mathsf{Exp}}_{\Psi, a \hookrightarrow \mathtt{petsm}(\rho; e_{\mathrm{parse}})} \mathtt{defref}[a] \,@\, \rho} \tag{5.31a}$$

$$\frac{\Omega, t :: \mathsf{Type} \vdash^{\mathsf{Exp}}_{\Psi} \epsilon \, @ \, \rho}{\Omega \vdash^{\mathsf{Exp}}_{\Psi} \mathtt{abstype}(t.\epsilon) \, @ \, \mathtt{alltypes}(t.\rho)} \tag{5.31b}$$

$$\frac{\Omega \vdash \sigma \, \mathsf{sig} \qquad \Omega, X : \sigma \vdash^{\mathsf{Exp}}_{\Psi} \epsilon \, @ \, \rho}{\Omega \vdash^{\mathsf{Exp}}_{\Psi} \mathtt{absmod}\{\sigma\}(X.\epsilon) \, @ \, \mathtt{allmods}\{\sigma\}(X.\rho)} \tag{5.31c}$$

$$\frac{\Omega \vdash^{\mathsf{Exp}}_{\Psi} \epsilon \, @ \, \mathtt{alltypes}(t.\rho) \qquad \Omega \vdash \tau \, \mathsf{type}}{\Omega \vdash^{\mathsf{Exp}}_{\Psi} \mathtt{aptype}\{\tau\}(\epsilon) \, @ \, [\tau/t]\rho} \tag{5.31d}$$

$$\frac{\Omega \vdash^{\mathsf{Exp}}_{\Psi} \epsilon \, @ \, \mathtt{allmods}\{\sigma\}(X'.\rho) \qquad \Omega \vdash X : \sigma}{\Omega \vdash^{\mathsf{Exp}}_{\Psi} \mathtt{apmod}\{X\}(\epsilon) \, @ \, [X/X']\rho} \tag{5.31e}$$

ppTSM Expression Typing

$$\frac{}{\Omega \vdash^{\mathsf{Pat}}_{\Phi, a \hookrightarrow \mathsf{pptsm}(\rho; e_{\mathrm{parse}})} \mathtt{defref}[a] \, @ \, \rho} \tag{5.32a}$$

$$\frac{\Omega, t :: \mathsf{Type} \vdash^{\mathsf{Pat}}_{\Phi} \epsilon \, @ \, \rho}{\Omega \vdash^{\mathsf{Pat}}_{\Phi} \mathtt{abstype}(t.\epsilon) \, @ \, \mathtt{alltypes}(t.\rho)} \tag{5.32b}$$

$$\frac{\Omega \vdash \sigma \, \mathsf{sig} \qquad \Omega, X : \sigma \vdash^{\mathsf{Pat}}_{\Phi} \epsilon \, @ \, \rho}{\Omega \vdash^{\mathsf{Pat}}_{\Phi} \mathtt{absmod}\{\sigma\}(X.\epsilon) \, @ \, \mathtt{allmods}\{\sigma\}(X.\rho)} \tag{5.32c}$$

$$\frac{\Omega \vdash^{\mathsf{Pat}}_{\Phi} \epsilon \, @ \, \mathtt{alltypes}(t.\rho) \qquad \Omega \vdash \tau \, \mathsf{type}}{\Omega \vdash^{\mathsf{Pat}}_{\Phi} \mathtt{aptype}\{\tau\}(\epsilon) \, @ \, [\tau/t]\rho} \tag{5.32d}$$

$$\frac{\Omega \vdash^{\mathsf{Pat}}_{\Phi} \epsilon \, @ \, \mathtt{allmods}\{\sigma\}(X'.\rho) \qquad \Omega \vdash X : \sigma}{\Omega \vdash^{\mathsf{Pat}}_{\Phi} \mathtt{apmod}\{X\}(\epsilon) \, @ \, [X/X']\rho} \tag{5.32e}$$

peTSM Expression Expansion

$$\frac{\Omega \vdash^{\mathsf{Exp}}_{\Psi} \epsilon \, @ \, \rho}{\langle \mathcal{D}; \mathcal{G}; \mathcal{M}; \Omega \rangle \vdash^{\mathsf{Exp}}_{\langle \mathcal{A}, \hat{a} \hookrightarrow \epsilon; \Psi; \mathcal{I} \rangle} \mathtt{bindref}[\hat{a}] \leadsto \epsilon \, @ \, \rho} \tag{5.33a}$$

$$\frac{\hat{\Omega}, \hat{t} \leadsto t :: \mathsf{Type} \vdash^{\mathsf{Exp}}_{\hat{\Psi}} \hat{\epsilon} \leadsto \epsilon \, @ \, \rho}{\hat{\Omega} \vdash^{\mathsf{Exp}}_{\hat{\Psi}} \mathtt{uabstype}(\hat{t}.\hat{\epsilon}) \leadsto \mathtt{abstype}(t.\epsilon) \, @ \, \mathtt{alltypes}(t.\rho)} \tag{5.33b}$$

$$\frac{\hat{\Omega} \vdash \hat{\sigma} \leadsto \sigma \, \mathsf{sig} \qquad \hat{\Omega}, \hat{X} \leadsto X : \sigma \vdash^{\mathsf{Exp}}_{\hat{\Psi}} \hat{\epsilon} \leadsto \epsilon \, @ \, \rho}{\hat{\Omega} \vdash^{\mathsf{Exp}}_{\hat{\Psi}} \mathtt{uabsmod}\{\hat{\sigma}\}(\hat{X}.\hat{\epsilon}) \leadsto \mathtt{absmod}\{\sigma\}(X.\epsilon) \, @ \, \mathtt{allmods}\{\sigma\}(X.\rho)} \tag{5.33c}$$

$$\frac{\hat{\Omega} \vdash^{\mathsf{Exp}}_{\hat{\Psi}} \hat{\epsilon} \leadsto \epsilon \, @ \, \mathtt{alltypes}(t.\rho) \qquad \hat{\Omega} \vdash \hat{\tau} \leadsto \tau \, \mathsf{type}}{\hat{\Omega} \vdash^{\mathsf{Exp}}_{\hat{\Psi}} \mathtt{uaptype}\{\hat{\tau}\}(\hat{\epsilon}) \leadsto \mathtt{aptype}\{\tau\}(\epsilon) \, @ \, [\tau/t]\rho} \tag{5.33d}$$

$$\frac{\hat{\Omega} \vdash^{\mathsf{Exp}}_{\hat{\Psi}} \hat{\epsilon} \rightsquigarrow \epsilon @ \mathtt{allmods}\{\sigma\}(X'.\rho) \qquad \hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{X} \rightsquigarrow X \Leftarrow \sigma}{\hat{\Omega} \vdash^{\mathsf{Exp}}_{\hat{\Psi}} \mathtt{uapmod}\{\hat{X}\}(\hat{\epsilon}) \rightsquigarrow \mathtt{apmod}\{X\}(\epsilon) @ [X/X']\rho} \tag{5.33e}$$

ppTSM expression expansion

$$\frac{\Omega \vdash^{\mathsf{Pat}}_{\Phi} \epsilon @ \rho}{\langle \mathcal{D}; \mathcal{G}; \mathcal{M}; \Omega \rangle \vdash^{\mathsf{Pat}}_{\langle \mathcal{A}, \hat{a} \hookrightarrow \epsilon; \Phi; \mathcal{I} \rangle} \mathtt{bindref}[\hat{a}] \rightsquigarrow \epsilon @ \rho} \tag{5.34a}$$

$$\frac{\hat{\Omega}, \hat{t} \rightsquigarrow t :: \mathsf{Type} \vdash^{\mathsf{Pat}}_{\hat{\Phi}} \hat{\epsilon} \rightsquigarrow \epsilon @ \rho}{\hat{\Omega} \vdash^{\mathsf{Pat}}_{\hat{\Phi}} \mathtt{uabstype}(\hat{t}.\hat{\epsilon}) \rightsquigarrow \mathtt{abstype}(t.\epsilon) @ \mathtt{alltypes}(t.\rho)} \tag{5.34b}$$

$$\frac{\hat{\Omega} \vdash \hat{\sigma} \rightsquigarrow \sigma\ \mathsf{sig} \qquad \hat{\Omega}, \hat{X} \rightsquigarrow X : \sigma \vdash^{\mathsf{Pat}}_{\hat{\Phi}} \hat{\epsilon} \rightsquigarrow \epsilon @ \rho}{\hat{\Omega} \vdash^{\mathsf{Pat}}_{\hat{\Phi}} \mathtt{uabsmod}\{\hat{\sigma}\}(\hat{X}.\hat{\epsilon}) \rightsquigarrow \mathtt{absmod}\{\sigma\}(X.\epsilon) @ \mathtt{allmods}\{\sigma\}(X.\rho)} \tag{5.34c}$$

$$\frac{\hat{\Omega} \vdash^{\mathsf{Pat}}_{\hat{\Phi}} \hat{\epsilon} \rightsquigarrow \epsilon @ \mathtt{alltypes}(t.\rho) \qquad \hat{\Omega} \vdash \hat{\tau} \rightsquigarrow \tau\ \mathsf{type}}{\hat{\Omega} \vdash^{\mathsf{Pat}}_{\hat{\Phi}} \mathtt{uaptype}\{\hat{\tau}\}(\hat{\epsilon}) \rightsquigarrow \mathtt{aptype}\{\tau\}(\epsilon) @ [\tau/t]\rho} \tag{5.34d}$$

$$\frac{\hat{\Omega} \vdash^{\mathsf{Pat}}_{\hat{\Phi}} \hat{\epsilon} \rightsquigarrow \epsilon @ \mathtt{allmods}\{\sigma\}(X'.\rho) \qquad \hat{\Omega} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{X} \rightsquigarrow X \Leftarrow \sigma}{\hat{\Omega} \vdash^{\mathsf{Pat}}_{\hat{\Phi}} \mathtt{uapmod}\{\hat{X}\}(\hat{\epsilon}) \rightsquigarrow \mathtt{apmod}\{X\}(\epsilon) @ [X/X']\rho} \tag{5.34e}$$

peTSM expression normal forms

$$\frac{}{\Omega \vdash^{\mathsf{Exp}}_{\Psi, a \hookrightarrow \mathtt{petsm}(\rho; e_{\mathrm{parse}})} \mathtt{defref}[a]\ \mathsf{normal}} \tag{5.35a}$$

$$\frac{}{\Omega \vdash^{\mathsf{Exp}}_{\Psi} \mathtt{abstype}(t.\epsilon)\ \mathsf{normal}} \tag{5.35b}$$

$$\frac{}{\Omega \vdash^{\mathsf{Exp}}_{\Psi} \mathtt{absmod}\{\sigma\}(X.\epsilon)\ \mathsf{normal}} \tag{5.35c}$$

$$\frac{\epsilon \neq \mathtt{abstype}(t.\epsilon') \qquad \Omega \vdash^{\mathsf{Exp}}_{\Psi} \epsilon\ \mathsf{normal}}{\Omega \vdash^{\mathsf{Exp}}_{\Psi} \mathtt{aptype}\{\tau\}(\epsilon)\ \mathsf{normal}} \tag{5.35d}$$

$$\frac{\epsilon \neq \mathtt{absmod}\{\sigma\}(X'.\epsilon') \qquad \Omega \vdash^{\mathsf{Exp}}_{\Psi} \epsilon\ \mathsf{normal}}{\Omega \vdash^{\mathsf{Exp}}_{\Psi} \mathtt{apmod}\{X\}(\epsilon)\ \mathsf{normal}} \tag{5.35e}$$

peTSM transitions

$$\frac{\Omega \vdash^{\mathsf{Exp}}_{\Psi} \epsilon \mapsto \epsilon'}{\Omega \vdash^{\mathsf{Exp}}_{\Psi} \mathtt{aptype}\{\tau\}(\epsilon) \mapsto \mathtt{aptype}\{\tau\}(\epsilon')} \tag{5.36a}$$

$$\frac{}{\Omega \vdash^{\mathsf{Exp}}_{\Psi} \mathtt{aptype}\{\tau\}(\mathtt{abstype}(t.\epsilon)) \mapsto [\tau/t]\epsilon} \tag{5.36b}$$

$$\frac{\Omega \vdash^{\mathsf{Exp}}_{\Psi} \epsilon \mapsto \epsilon'}{\Omega \vdash^{\mathsf{Exp}}_{\Psi} \mathtt{apmod}\{X\}(\epsilon) \mapsto \mathtt{apmod}\{X\}(\epsilon')} \tag{5.36c}$$

$$\frac{}{\Omega \vdash^{\mathsf{Exp}}_{\Psi} \mathtt{apmod}\{X\}(\mathtt{absmod}\{\sigma\}(X'.\epsilon)) \mapsto [X/X']\epsilon} \tag{5.36d}$$

peTSM reflexive, transitive transitions

$$\frac{}{\Omega \vdash^{\mathsf{Exp}}_{\Psi} \epsilon \mapsto^{*} \epsilon} \tag{5.37a}$$

$$\frac{\Omega \vdash^{\mathsf{Exp}}_{\Psi} \epsilon \mapsto \epsilon'}{\Omega \vdash^{\mathsf{Exp}}_{\Psi} \epsilon \mapsto^{*} \epsilon'} \tag{5.37b}$$

$$\frac{\Omega \vdash^{\mathsf{Exp}}_{\Psi} \epsilon \mapsto^{*} \epsilon' \qquad \Omega \vdash^{\mathsf{Exp}}_{\Psi} \epsilon' \mapsto^{*} \epsilon''}{\Omega \vdash^{\mathsf{Exp}}_{\Psi} \epsilon \mapsto^{*} \epsilon''} \tag{5.37c}$$

peTSM normalization

$$\frac{\Omega \vdash^{\mathsf{Exp}}_{\Psi} \epsilon \mapsto^{*} \epsilon' \qquad \Omega \vdash^{\mathsf{Exp}}_{\Psi} \epsilon' \text{ normal}}{\Omega \vdash^{\mathsf{Exp}}_{\Psi} \epsilon \Downarrow \epsilon'} \tag{5.38}$$

TSM expression definition extraction

$$\mathsf{tsmdef}(\mathtt{defref}[a]) = a \tag{5.39a}$$
$$\mathsf{tsmdef}(\mathtt{abstype}(t.\epsilon)) = \mathsf{tsmdef}(\epsilon) \tag{5.39b}$$
$$\mathsf{tsmdef}(\mathtt{absmod}\{\sigma\}(X.\epsilon)) = \mathsf{tsmdef}(\epsilon) \tag{5.39c}$$
$$\mathsf{tsmdef}(\mathtt{aptype}\{\tau\}(\epsilon)) = \mathsf{tsmdef}(\epsilon) \tag{5.39d}$$
$$\mathsf{tsmdef}(\mathtt{apmod}\{X\}(\epsilon)) = \mathsf{tsmdef}(\epsilon) \tag{5.39e}$$

5.2.6 Syntax of Candidate Expansions

Figure 5.10 defines the syntax of candidate expansion types (or *ce-types*), $\grave{\tau}$, candidate expansion expressions (or *ce-expressions*), $\grave{e}$, candidate expansion rules (or *ce-rules*), $\grave{r}$, and candidate expansion patterns (or *ce-patterns*), $\grave{p}$. Candidate expansion terms are identified up to α-equivalence in the usual manner.

Each inner core form, except for the variable pattern form, maps onto a candidate expansion form. In particular:

- Each type form maps onto a ce-type form according to the metafunction $\mathcal{C}(\tau)$, defined in Sec. 3.2.8.

Sort		Operational Form	Stylized Form	Description
PCEExp	$\check{e}$ ::=	$\texttt{ceexp}(\grave{e})$	$\grave{e}$	ce-expression
		$\texttt{cebindtype}(t.\check{e})$	$\Lambda t.\check{e}$	type binding
		$\texttt{cebindmod}(X.\check{e})$	$\Lambda X.\check{e}$	module binding

Figure 5.8: Abstract syntax of parameterized candidate expansion expressions in miniVerse$_\forall$. Parameterized candidate expansion expressions are identified up to α-equivalence.

Sort		Operational Form	Stylized Form	Description
CEKind	$\check{\kappa}$ ::=	$\texttt{cedarr}(\check{\kappa};u.\check{\kappa})$	$(u :: \check{\kappa}) \rightarrow \check{\kappa}$	dependent function
		$\texttt{ceunit}$	$\langle\!\langle\,\rangle\!\rangle$	nullary product
		$\texttt{cedprod}(\check{\kappa};u.\check{\kappa})$	$(u :: \check{\kappa}) \times \check{\kappa}$	dependent product
		$\texttt{ceType}$	T	types
		$\texttt{ceS}(\tilde{\tau})$	$[=\tilde{\tau}]$	singleton
		$\texttt{cesplicedk}[m;n]$	$\texttt{splicedk}\langle m,n\rangle$	spliced
CECon	$\grave{c},\tilde{\tau}$ ::=	u	u	constructor variable
		t	t	type variable
		$\texttt{ceasc}\{\check{\kappa}\}(\grave{c})$	$\grave{c} :: \check{\kappa}$	ascription
		$\texttt{ceabs}(u.\grave{c})$	$\lambda u.\grave{c}$	abstraction
		$\texttt{ceapp}(\grave{c};\grave{c})$	$\grave{c}(\grave{c})$	application
		$\texttt{cetriv}$	$\langle\!\langle\,\rangle\!\rangle$	trivial
		$\texttt{cepair}(\grave{c};\grave{c})$	$\langle\!\langle \grave{c},\grave{c}\rangle\!\rangle$	pair
		$\texttt{ceprl}(\grave{c})$	$\grave{c} \cdot \texttt{l}$	left projection
		$\texttt{ceprr}(\grave{c})$	$\grave{c} \cdot \texttt{r}$	right projection
		$\texttt{ceparr}(\tilde{\tau};\tilde{\tau})$	$\tilde{\tau} \rightharpoonup \tilde{\tau}$	partial function
		$\texttt{ceall}\{\check{\kappa}\}(u.\tilde{\tau})$	$\forall(u :: \check{\kappa}).\tilde{\tau}$	polymorphic
		$\texttt{cerec}(t.\tilde{\tau})$	$\mu t.\tilde{\tau}$	recursive
		$\texttt{ceprod}[L](\{i \hookrightarrow \tilde{\tau}_i\}_{i\in L})$	$\langle\{i \hookrightarrow \tilde{\tau}_i\}_{i\in L}\rangle$	labeled product
		$\texttt{cesum}[L](\{i \hookrightarrow \tilde{\tau}_i\}_{i\in L})$	$[\{i \hookrightarrow \tilde{\tau}_i\}_{i\in L}]$	labeled sum
		$\texttt{cecon}(X)$	$X \cdot \texttt{c}$	constructor part
		$\texttt{cesplicedc}[m;n]$	$\texttt{splicedc}\langle m,n\rangle$	spliced

Figure 5.9: Syntax of candidate expansion kinds and constructors in miniVerse$_\forall$. Candidate expansion kinds and constructors are identified up to α-equivalence.

Sort			Operational Form	Stylized Form	Description
CEExp	$\grave{e}$	::=	x	x	variable
			$\mathtt{ceasc}\{\tilde{\tau}\}(\grave{e})$	$\grave{e}:\tilde{\tau}$	ascription
			$\mathtt{celetval}(\grave{e};x.\grave{e})$	$\mathtt{let\ val}\ x = \grave{e}\ \mathtt{in}\ \grave{e}$	value binding
			$\mathtt{ceanalam}(x.\grave{e})$	$\lambda x.\grave{e}$	abstraction (unannotated)
			$\mathtt{celam}\{\tilde{\tau}\}(x.\grave{e})$	$\lambda x{:}\tilde{\tau}.\grave{e}$	abstraction (annotated)
			$\mathtt{ceap}(\grave{e};\grave{e})$	$\grave{e}(\grave{e})$	application
			$\mathtt{ceclam}\{\check{\kappa}\}(u.\grave{e})$	$\Lambda u{::}\check{\kappa}.\grave{e}$	constructor abstraction
			$\mathtt{cecap}\{\hat{c}\}(\grave{e})$	$\grave{e}[\hat{c}]$	constructor application
			$\mathtt{cefold}(\grave{e})$	$\mathtt{fold}(\grave{e})$	fold
			$\mathtt{ceunfold}(\grave{e})$	$\mathtt{unfold}(\grave{e})$	unfold
			$\mathtt{cetpl}[L](\{i \hookrightarrow \grave{e}_i\}_{i \in L})$	$\langle\{i \hookrightarrow \grave{e}_i\}_{i \in L}\rangle$	labeled tuple
			$\mathtt{cepr}[\ell](\grave{e})$	$\grave{e}\cdot\ell$	projection
			$\mathtt{cein}[\ell](\grave{e})$	$\mathtt{inj}\ \ell\cdot\grave{e}$	injection
			$\mathtt{cematch}[n](\grave{e};\{\hat{r}_i\}_{1\le i\le n})$	$\mathtt{match}\ \grave{e}\ \{\hat{r}_i\}_{1\le i\le n}$	match
			$\mathtt{ceval}(X)$	$X\cdot\mathtt{v}$	value part
			$\mathtt{cesplicede}[m;n]$	$\mathtt{splicede}\langle m,n\rangle$	spliced
CERule	$\hat{r}$	::=	$\mathtt{cerule}(p.\grave{e})$	$p \Rightarrow \grave{e}$	rule
CEPat	$\grave{p}$	::=	$\mathtt{cewildp}$	$\_$	wildcard pattern
			$\mathtt{cefoldp}(p)$	$\mathtt{fold}(p)$	fold pattern
			$\mathtt{cetplp}[L](\{i \hookrightarrow \grave{p}_i\}_{i \in L})$	$\langle\{i \hookrightarrow \grave{p}_i\}_{i \in L}\rangle$	labeled tuple pattern
			$\mathtt{ceinp}[\ell](\grave{p})$	$\mathtt{inj}\ \ell\cdot\grave{p}$	injection pattern
			$\mathtt{cesplicedp}[m;n]$	$\mathtt{splicedp}\langle m,n\rangle$	spliced

Figure 5.10: Abstract syntax of candidate expansion types, expressions, rules and patterns in miniVerse$_\forall$. Candidate expansion terms are identified up to α-equivalence.

- Each expanded expression form maps onto a ce-expression form according to the metafunction $\mathcal{C}(e)$, defined as follows:

$$\mathcal{C}(x) = x$$
$$\mathcal{C}(\texttt{lam}\{\tau\}(x.e)) = \texttt{celam}\{\mathcal{C}(\tau)\}(x.\mathcal{C}(e))$$
$$\mathcal{C}(\texttt{ap}(e_1;e_2)) = \texttt{ceap}(\mathcal{C}(e_1);\mathcal{C}(e_2))$$
$$\mathcal{C}(\texttt{tlam}(t.e)) = \texttt{cetlam}(t.\mathcal{C}(e))$$
$$\mathcal{C}(\texttt{tap}\{\tau\}(e)) = \texttt{cetap}\{\mathcal{C}(\tau)\}(\mathcal{C}(e))$$
$$\mathcal{C}(\texttt{fold}\{t.\tau\}(e)) = \texttt{ceasc}\{\texttt{cerec}(t.\mathcal{C}(\tau))\}(\texttt{cefold}(\mathcal{C}(e)))$$
$$\mathcal{C}(\texttt{unfold}(e)) = \texttt{ceunfold}(\mathcal{C}(e))$$
$$\mathcal{C}(\texttt{tpl}[L](\{i \hookrightarrow e_i\}_{i \in L})) = \texttt{cetpl}[L](\{i \hookrightarrow \mathcal{C}(e_i)\}_{i \in L})$$
$$\mathcal{C}(\texttt{in}[L][\ell]\{\{i \hookrightarrow \tau_i\}_{i \in L}\}(e)) = \texttt{ceasc}\{\texttt{cesum}[L](\{i \hookrightarrow \mathcal{C}(\tau_i)\}_{i \in L})\}(\texttt{cein}[\ell](\mathcal{C}(e)))$$
$$\mathcal{C}(\texttt{match}[n]\{\tau\}(e;\{r_i\}_{1 \leq i \leq n})) = \texttt{ceasc}\{\mathcal{C}(\tau)\}(\texttt{cematch}[n](\mathcal{C}(e);\{\mathcal{C}(r_i)\}_{1 \leq i \leq n}))$$

- The expanded rule form maps onto the ce-rule form according to the metafunction $\mathcal{C}(r)$, defined as follows:

$$\mathcal{C}(\texttt{rule}(p.e)) = \texttt{cerule}(p.\mathcal{C}(e))$$

- Each expanded pattern form, except for the variable pattern form, maps onto a ce-pattern form according to the metafunction $\mathcal{C}(p)$, defined in Sec. 4.2.8.

There are three other candidate expansion forms: a ce-type form for *references to spliced unexpanded types*, $\texttt{cesplicedt}[m;n]$, a ce-expression form for *references to spliced unexpanded expressions*, $\texttt{cesplicede}[m;n]$, and a ce-pattern form for *references to spliced unexpanded patterns*, $\texttt{cesplicedp}[m;n]$.

5.2.7 Candidate Expansion Validation

Judgement Form	Description
$\Omega_{\text{app}} \vdash^{\textsf{Exp}}_{\Psi} \check{e} \leftrightarrow_\epsilon \grave{e} ? \rho \dashv \omega : \Omega_{\text{params}}$	$\check{e}$ has deparameterization $\grave{e} ? \rho$ when generated by ϵ, with substitution ω for variables tracked by Ω_{params}
$\Omega_{\text{app}} \vdash^{\textsf{Pat}}_{\Phi} \cdot \leftrightarrow_\epsilon \cdot ? \rho \dashv \omega : \Omega_{\text{params}}$	Any ce-pattern generated by ϵ has deparameterization $\cdot ? \rho$ with substitution ω for variables tracked by Ω_{params}

Candidate Expansion Expression Deparameterization

$$\frac{}{\Omega_{\text{app}} \vdash^{\textsf{Exp}}_{\Psi,a \hookrightarrow \texttt{petsm}(\rho;e_{\text{parse}})} \texttt{ceexp}(\grave{e}) \leftrightarrow_{\texttt{defref}[a]} \grave{e} ? \rho \dashv \varnothing : \varnothing} \quad (5.40\text{a})$$

$$\frac{\Omega_{\mathrm{app}} \vdash^{\mathsf{Exp}}_{\Psi} \breve{e} \looparrowright_{\epsilon} \grave{e} \,?\, \mathtt{alltypes}(t.\rho) \dashv \omega : \Omega \qquad t \notin \mathrm{dom}(\Omega_{\mathrm{app}})}{\Omega_{\mathrm{app}} \vdash^{\mathsf{Exp}}_{\Psi} \mathtt{cebindtype}(t.\breve{e}) \looparrowright_{\mathtt{aptype}\{\tau\}(\epsilon)} \grave{e} \,?\, \rho \dashv \omega, \tau/t : \Omega, t :: \mathtt{Type}} \tag{5.40b}$$

$$\frac{\Omega_{\mathrm{app}} \vdash^{\mathsf{Exp}}_{\Psi} \breve{e} \looparrowright_{\epsilon} \grave{e} \,?\, \mathtt{allmods}\{\sigma\}(X.\rho) \dashv \omega : \Omega \qquad X \notin \mathrm{dom}(\Omega_{\mathrm{app}})}{\Omega_{\mathrm{app}} \vdash^{\mathsf{Exp}}_{\Psi} \mathtt{cebindmod}(X.\breve{e}) \looparrowright_{\mathtt{apmod}\{X'\}(\epsilon)} \grave{e} \,?\, \rho \dashv \omega, X'/X : \Omega, X : \sigma} \tag{5.40c}$$

Candidate Expansion Pattern Deparameterization

$$\frac{}{\Omega_{\mathrm{app}} \vdash^{\mathsf{Pat}}_{\Phi,a \hookrightarrow \mathtt{pptsm}(\rho; e_{\mathrm{parse}})} \cdot \looparrowright_{\mathtt{defref}[a]} \cdot \,?\, \rho \dashv \varnothing : \varnothing} \tag{5.41a}$$

$$\frac{\Omega_{\mathrm{app}} \vdash^{\mathsf{Pat}}_{\Phi} \cdot \looparrowright_{\epsilon} \cdot \,?\, \mathtt{alltypes}(t.\rho) \dashv \omega : \Omega \qquad t \notin \mathrm{dom}(\Omega_{\mathrm{app}})}{\Omega_{\mathrm{app}} \vdash^{\mathsf{Pat}}_{\Phi} \cdot \looparrowright_{\mathtt{aptype}\{\tau\}(\epsilon)} \cdot \,?\, \rho \dashv \omega, \tau/t : \Omega, t :: \mathtt{Type}} \tag{5.41b}$$

$$\frac{\Omega_{\mathrm{app}} \vdash^{\mathsf{Pat}}_{\Phi} \cdot \looparrowright_{\epsilon} \cdot \,?\, \mathtt{allmods}\{\sigma\}(X.\rho) \dashv \omega : \Omega \qquad X \notin \mathrm{dom}(\Omega_{\mathrm{app}})}{\Omega_{\mathrm{app}} \vdash^{\mathsf{Pat}}_{\Phi} \cdot \looparrowright_{\mathtt{apmod}\{X'\}(\epsilon)} \cdot \,?\, \rho \dashv \omega, X'/X : \Omega, X : \sigma} \tag{5.41c}$$

The *bidirectional candidate expansion validation judgements* validate ce-terms and simultaneously generate their final expansions.

Judgement Form	Description
$\Omega \vdash^{\mathbb{C}} \check{\kappa} \rightsquigarrow \kappa$ kind	$\check{\kappa}$ is well-formed and has expansion κ
$\Omega \vdash^{\mathbb{C}} \check{c} \rightsquigarrow c \Rightarrow \kappa$	$\check{c}$ has expansion c and synthesizes kind κ
$\Omega \vdash^{\mathbb{C}} \check{c} \rightsquigarrow c \Leftarrow \kappa$	$\check{c}$ has expansion c when analyzed against kind κ

Judgement Form	Description
$\Omega \vdash^{\mathbb{C}} \check{\tau} \rightsquigarrow \tau$ type	$\check{\tau}$ has expansion τ
$\Omega \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Rightarrow \tau$	$\grave{e}$ has expansion e and synthesizes type τ
$\Omega \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Leftarrow \tau$	$\grave{e}$ has expansion e when analyzed against type τ
$\Omega \vdash^{\mathbb{E}} \grave{r} \rightsquigarrow r \Rightarrow \tau \mapsto \tau'$	$\grave{r}$ has expansion r and takes values of type τ to values of synthesized type τ'
$\Omega \vdash^{\mathbb{E}} \grave{r} \rightsquigarrow r \Leftarrow \tau \mapsto \tau'$	$\grave{r}$ has expansion r and takes values of type τ to values of type τ' when τ' is provided for analysis
$\Omega_{\mathrm{param}} \vdash^{\mathbb{P}} \grave{p} \rightsquigarrow p : \tau \dashv\!|^{\hat{\Omega}}$	$\grave{p}$ expands to p and matches values of type τ generating assumptions $\hat{Y}$

Expression splicing scenes, $\mathbb{E}$, are of the form $\hat{\Omega}; \hat{\Psi}; \hat{\Phi}; b$, *constructor splicing scenes*, $\mathbb{C}$, are of the form $\hat{\Omega}; b$, and *pattern splicing scenes*, $\mathbb{P}$, are of the form $\hat{\Omega}; \hat{\Phi}; b$. Their purpose is to "remember", during candidate expansion validation, the contexts, TSM environments and literal bodies from the TSM application site (cf. Rules (4.4m) and (4.6f)), because these are necessary to validate references to spliced terms. We write $\mathrm{cs}(\mathbb{E})$ for the constructor splicing scene constructed by dropping the TSM contexts from $\mathbb{E}$:

$$\mathrm{cs}(\hat{\Omega}; \hat{\Psi}; \hat{\Phi}; b) = \hat{\Omega}; b$$

Candidate Expansion Kind and Constructor Validation

ce-kind validation

$$\frac{\Omega \vdash^{\mathbb{C}} \check{\kappa}_1 \rightsquigarrow \kappa_1 \text{ kind} \qquad \Omega, u :: \kappa_1 \vdash^{\mathbb{C}} \check{\kappa}_2 \rightsquigarrow \kappa_2 \text{ kind}}{\Omega \vdash^{\mathbb{C}} \text{cedarr}(\check{\kappa}_1; u.\check{\kappa}_2) \rightsquigarrow \text{darr}(\kappa_1; u.\kappa_2) \text{ kind}} \tag{5.42a}$$

$$\frac{}{\Omega \vdash^{\mathbb{C}} \text{ceunit} \rightsquigarrow \text{unit kind}} \tag{5.42b}$$

$$\frac{\Omega \vdash^{\mathbb{C}} \check{\kappa}_1 \rightsquigarrow \kappa_1 \text{ kind} \qquad \Omega, u :: \kappa_1 \vdash^{\mathbb{C}} \check{\kappa}_2 \rightsquigarrow \kappa_2 \text{ kind}}{\Omega \vdash^{\mathbb{C}} \text{cedprod}(\check{\kappa}_1; u.\check{\kappa}_2) \rightsquigarrow \text{dprod}(\kappa_1; u.\kappa_2) \text{ kind}} \tag{5.42c}$$

$$\frac{}{\Omega \vdash^{\mathbb{C}} \text{ceType} \rightsquigarrow \text{Type kind}} \tag{5.42d}$$

$$\frac{\Omega \vdash^{\mathbb{C}} \check{\tau} \rightsquigarrow \tau \Leftarrow \text{Type}}{\Omega \vdash^{\mathbb{C}} \text{ceS}(\check{\tau}) \rightsquigarrow \text{S}(\tau) \text{ kind}} \tag{5.42e}$$

$$\frac{\begin{array}{c} \text{parseUKind}(\text{subseq}(b; m; n)) = \hat{\kappa} \qquad \hat{\Omega} \vdash \hat{\kappa} \rightsquigarrow \kappa \text{ kind} \\ \hat{\Omega} = \langle \mathcal{D}; \mathcal{G}; \mathcal{M}; \Omega_{\text{app}} \rangle \qquad \text{dom}(\Omega) \cap \text{dom}(\Omega_{\text{app}}) = \varnothing \end{array}}{\Omega \vdash^{\hat{\Omega}; b} \text{cesplicedk}[m; n] \rightsquigarrow \kappa \text{ kind}} \tag{5.42f}$$

Synthetic ce-constructor validation

$$\frac{}{\Omega, u :: \kappa \vdash^{\mathbb{C}} u \rightsquigarrow u \Rightarrow \kappa} \tag{5.43a}$$

$$\frac{\Omega \vdash^{\mathbb{C}} \check{\kappa} \rightsquigarrow \kappa \text{ kind} \qquad \Omega \vdash^{\mathbb{C}} \check{c} \rightsquigarrow c \Leftarrow \kappa}{\Omega \vdash^{\mathbb{C}} \text{ceasc}\{\check{\kappa}\}(\check{c}) \rightsquigarrow c \Rightarrow \kappa} \tag{5.43b}$$

$$\frac{\Omega \vdash^{\mathbb{C}} \check{c}_1 \rightsquigarrow c_1 \Rightarrow \text{darr}(\kappa_2; u.\kappa) \qquad \Omega \vdash^{\mathbb{C}} \check{c}_2 \rightsquigarrow c_2 \Rightarrow \kappa_2}{\Omega \vdash^{\mathbb{C}} \text{ceapp}(\check{c}_1; \check{c}_2) \rightsquigarrow \text{app}(c_1; c_2) \Rightarrow [c_1/u]\kappa} \tag{5.43c}$$

$$\frac{}{\Omega \vdash^{\mathbb{C}} \text{cetriv} \rightsquigarrow \text{triv} \Rightarrow \text{unit}} \tag{5.43d}$$

$$\frac{\Omega \vdash^{\mathbb{C}} \check{c} \rightsquigarrow c \Rightarrow \text{dprod}(\kappa_1; u.\kappa_2)}{\Omega \vdash^{\mathbb{C}} \text{ceprl}(\check{c}) \rightsquigarrow \text{prl}(c) \Rightarrow \kappa_1} \tag{5.43e}$$

$$\frac{\Omega \vdash^{\mathbb{C}} \check{c} \rightsquigarrow c \Rightarrow \text{dprod}(\kappa_1; u.\kappa_2)}{\Omega \vdash^{\mathbb{C}} \text{ceprr}(\check{c}) \rightsquigarrow \text{prr}(c) \Rightarrow [\text{prl}(c)/u]\kappa_2} \tag{5.43f}$$

$$\frac{\Omega \vdash^{\mathbb{C}} \check{\tau}_1 \rightsquigarrow \tau_1 \Leftarrow \text{Type} \qquad \Omega \vdash^{\mathbb{C}} \check{\tau}_2 \rightsquigarrow \tau_2 \Leftarrow \text{Type}}{\Omega \vdash^{\mathbb{C}} \text{ceparr}(\check{\tau}_1; \check{\tau}_2) \rightsquigarrow \text{parr}(\tau_1; \tau_2) \Rightarrow \text{Type}} \tag{5.43g}$$

$$\frac{\Omega \vdash^{\mathbb{C}} \check{\kappa} \rightsquigarrow \kappa \text{ kind} \qquad \Omega, u :: \kappa \vdash^{\mathbb{C}} \check{\tau} \rightsquigarrow \tau \Leftarrow \text{Type}}{\Omega \vdash^{\mathbb{C}} \text{ceall}\{\check{\kappa}\}(u.\check{\tau}) \rightsquigarrow \text{all}\{\kappa\}(u.\tau) \Rightarrow \text{Type}} \tag{5.43h}$$

$$\frac{\Omega, t :: \mathtt{Type} \vdash^{\mathbb{C}} \grave{\tau} \rightsquigarrow \tau \Leftarrow \mathtt{Type}}{\Omega \vdash^{\mathbb{C}} \mathtt{cerec}(t.\grave{\tau}) \rightsquigarrow \mathtt{rec}(t.\tau) \Rightarrow \mathtt{Type}} \tag{5.43i}$$

$$\frac{\{\Omega \vdash^{\mathbb{C}} \grave{\tau}_i \rightsquigarrow \tau_i \Leftarrow \mathtt{Type}\}_{1 \leq i \leq n}}{\Omega \vdash^{\mathbb{C}} \mathtt{uprod}[L](\{i \hookrightarrow \grave{\tau}_i\}_{i \in L}) \rightsquigarrow \mathtt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L}) \Rightarrow \mathtt{Type}} \tag{5.43j}$$

$$\frac{\{\Omega \vdash^{\mathbb{C}} \grave{\tau}_i \rightsquigarrow \tau_i \Leftarrow \mathtt{Type}\}_{1 \leq i \leq n}}{\Omega \vdash^{\mathbb{C}} \mathtt{cesum}[L](\{i \hookrightarrow \grave{\tau}_i\}_{i \in L}) \rightsquigarrow \mathtt{sum}[L](\{i \hookrightarrow \tau_i\}_{i \in L}) \Rightarrow \mathtt{Type}} \tag{5.43k}$$

$$\frac{}{\Omega, X : \mathtt{sig}\{\kappa\}(u.\tau) \vdash^{\mathbb{C}} \mathtt{cecon}(X) \rightsquigarrow \mathtt{con}(X) \Rightarrow \kappa} \tag{5.43l}$$

$$\frac{\begin{array}{c} \mathrm{parseUCon}(\mathrm{subseq}(b; m; n)) = \hat{c} \qquad \hat{\Omega} \vdash \hat{c} \rightsquigarrow c \Rightarrow \kappa \\ \hat{\Omega} = \langle \mathcal{D}; \mathcal{G}; \mathcal{M}; \Omega_{\mathrm{app}} \rangle \qquad \mathrm{dom}(\Omega) \cap \mathrm{dom}(\Omega_{\mathrm{app}}) = \varnothing \end{array}}{\Omega \vdash^{\hat{\Omega}; b} \mathtt{cesplicedc}[m; n] \rightsquigarrow c \Rightarrow \kappa} \tag{5.43m}$$

Analytic constructor expansion

$$\frac{\Omega \vdash^{\mathbb{C}} \grave{c} \rightsquigarrow c \Rightarrow \kappa_1 \qquad \Omega \vdash \kappa_1 <:: \kappa_2}{\Omega \vdash^{\mathbb{C}} \grave{c} \rightsquigarrow c \Leftarrow \kappa_2} \tag{5.44a}$$

$$\frac{\hat{\Omega} \vdash \grave{c} \rightsquigarrow c \Leftarrow \mathtt{Type}}{\hat{\Omega} \vdash \grave{c} \rightsquigarrow c \Leftarrow \mathtt{S}(c)} \tag{5.44b}$$

$$\frac{\Omega, u :: \kappa_1 \vdash^{\mathbb{C}} \grave{c}_2 \rightsquigarrow c_2 \Leftarrow \kappa_2}{\Omega \vdash^{\mathbb{C}} \mathtt{ceabs}(u.\grave{c}_2) \rightsquigarrow \mathtt{abs}(u.c_2) \Leftarrow \mathtt{darr}(\kappa_1; u.\kappa_2)} \tag{5.44c}$$

$$\frac{\Omega \vdash^{\mathbb{C}} \grave{c}_1 \rightsquigarrow c_1 \Leftarrow \kappa_1 \qquad \Omega \vdash^{\mathbb{C}} \grave{c}_2 \rightsquigarrow c_2 \Leftarrow [c_1/u]\kappa_2}{\Omega \vdash^{\mathbb{C}} \mathtt{cepair}(\grave{c}_1; \grave{c}_2) \rightsquigarrow \mathtt{pair}(c_1; c_2) \Leftarrow \mathtt{dprod}(\kappa_1; u.\kappa_2)} \tag{5.44d}$$

$$\frac{\begin{array}{c} \mathrm{parseUCon}(\mathrm{subseq}(b; m; n)) = \hat{c} \qquad \hat{\Omega} \vdash \hat{c} \rightsquigarrow c \Leftarrow \kappa \\ \hat{\Omega} = \langle \mathcal{D}; \mathcal{G}; \mathcal{M}; \Omega_{\mathrm{app}} \rangle \qquad \mathrm{dom}(\Omega) \cap \mathrm{dom}(\Omega_{\mathrm{app}}) = \varnothing \end{array}}{\Omega \vdash^{\hat{\Omega}; b} \mathtt{cesplicedc}[m; n] \rightsquigarrow c \Leftarrow \kappa} \tag{5.44e}$$

Bidirectional Candidate Expansion Expression Validation

Like the bidirectionally typed expression expansion judgements, the bidirectional ce-expression validation judgements distinguish type synthesis from type analysis. The *synthetic ce-expression validation judgement*, $\Delta \; \Gamma \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Rightarrow \tau$, and the *analytic ce-expression validation judgement*, $\Delta \; \Gamma \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Leftarrow \tau$, are defined mutually inductively with Rules (7.1) and Rules (7.2) by Rules (7.6) and Rules (7.7), respectively, as follows.

$$\frac{\Omega \vdash^{\mathbb{C}} \grave{\tau} \rightsquigarrow \tau \Leftarrow \mathtt{Type}}{\Omega \vdash^{\mathbb{C}} \grave{\tau} \rightsquigarrow \tau \; \mathtt{type}} \tag{5.45}$$

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Type Synthesis Synthetic ce-expression validation is governed by the following rules.

$$\frac{}{\Omega, x : \tau \vdash^{\mathbb{E}} x \rightsquigarrow x \Rightarrow \tau} \tag{5.46a}$$

$$\frac{\Omega \vdash^{\mathsf{cs}(\mathbb{E})} \grave{\tau} \rightsquigarrow \tau \text{ type} \qquad \Omega \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Leftarrow \tau}{\Omega \vdash^{\mathbb{E}} \mathtt{ceasc}\{\grave{\tau}\}(\grave{e}) \rightsquigarrow e \Rightarrow \tau} \tag{5.46b}$$

$$\frac{\Omega \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Rightarrow \tau \qquad \Omega, x : \tau \vdash^{\mathbb{E}} \grave{e}' \rightsquigarrow e' \Rightarrow \tau'}{\Omega \vdash^{\mathbb{E}} \mathtt{celetval}(\grave{e}; x.\grave{e}') \rightsquigarrow \mathtt{ap}(\mathtt{lam}\{\tau\}(x.e'); e) \Rightarrow \tau'} \tag{5.46c}$$

$$\frac{\Omega \vdash^{\mathsf{cs}(\mathbb{E})} \grave{\tau}_1 \rightsquigarrow \tau_1 \text{ type} \qquad \Omega, x : \tau_1 \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Rightarrow \tau_2}{\Omega \vdash^{\mathbb{E}} \mathtt{celam}\{\grave{\tau}_1\}(x.\grave{e}) \rightsquigarrow \mathtt{lam}\{\tau_1\}(x.e) \Rightarrow \mathtt{parr}(\tau_1; \tau_2)} \tag{5.46d}$$

$$\frac{\Omega \vdash^{\mathbb{E}} \grave{e}_1 \rightsquigarrow e_1 \Rightarrow \mathtt{parr}(\tau_2; \tau) \qquad \Omega \vdash^{\mathbb{E}} \grave{e}_2 \rightsquigarrow e_2 \Leftarrow \tau_2}{\Omega \vdash^{\mathbb{E}} \mathtt{ceap}(\grave{e}_1; \grave{e}_2) \rightsquigarrow \mathtt{ap}(e_1; e_2) \Rightarrow \tau} \tag{5.46e}$$

$$\frac{\Omega \vdash^{\mathsf{cs}(\mathbb{E})} \grave{\kappa} \rightsquigarrow \kappa \text{ kind} \qquad \Omega, u :: \kappa \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Rightarrow \tau}{\Delta \, \Gamma \vdash^{\mathbb{E}} \mathtt{ceclam}\{\grave{\kappa}\}(u.\grave{e}) \rightsquigarrow \mathtt{clam}\{\kappa\}(u.e) \Rightarrow \mathtt{all}\{\kappa\}(u.\tau)} \tag{5.46f}$$

$$\frac{\Omega \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Rightarrow \mathtt{all}\{\kappa\}(u.\tau) \qquad \Omega \vdash^{\mathsf{cs}(\mathbb{E})} \grave{c} \rightsquigarrow c \Leftarrow \kappa}{\Omega \vdash^{\mathbb{E}} \mathtt{cecap}\{\grave{c}\}(\grave{e}) \rightsquigarrow \mathtt{cap}\{c\}(e) \Rightarrow [c/u]\tau} \tag{5.46g}$$

$$\frac{\Omega \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Rightarrow \mathtt{rec}(t.\tau)}{\Omega \vdash^{\mathbb{E}} \mathtt{ceunfold}(\grave{e}) \rightsquigarrow \mathtt{unfold}(e) \Rightarrow [\mathtt{rec}(t.\tau)/t]\tau} \tag{5.46h}$$

$$\frac{\begin{array}{cc} \grave{e} = \mathtt{cetpl}[L](\{i \hookrightarrow \grave{e}_i\}_{i \in L}) & e = \mathtt{tpl}[L](\{i \hookrightarrow e_i\}_{i \in L}) \\ \{\Omega \vdash^{\mathbb{E}} \grave{e}_i \rightsquigarrow e_i \Rightarrow \tau_i\}_{i \in L} \end{array}}{\Omega \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Rightarrow \mathtt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L})} \tag{5.46i}$$

$$\frac{\Omega \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Rightarrow \mathtt{prod}[L, \ell](\{i \hookrightarrow \tau_i\}_{i \in L}; \ell \hookrightarrow \tau)}{\Omega \vdash^{\mathbb{E}} \mathtt{cepr}[\ell](\grave{e}) \rightsquigarrow \mathtt{pr}[\ell](e) \Rightarrow \tau} \tag{5.46j}$$

$$\frac{n > 0 \qquad \Omega \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Rightarrow \tau \qquad \{\Omega \vdash^{\mathbb{E}} \grave{r}_i \rightsquigarrow r_i \Rightarrow \tau \Mapsto \tau'\}_{1 \leq i \leq n}}{\Omega \vdash^{\mathbb{E}} \mathtt{cematch}[n](\grave{e}; \{\grave{r}_i\}_{1 \leq i \leq n}) \rightsquigarrow \mathtt{match}[n]\{\tau'\}(e; \{r_i\}_{1 \leq i \leq n}) \Rightarrow \tau'} \tag{5.46k}$$

$$\frac{}{\Omega, X : \mathtt{sig}\{\kappa\}(u.\tau) \vdash^{\mathbb{E}} \mathtt{ceval}(X) \rightsquigarrow \mathtt{val}(X) \Rightarrow [\mathtt{con}(X)/u]\tau} \tag{5.46l}$$

$$\frac{\begin{array}{cc} \mathtt{parseUExp}(\mathtt{subseq}(b; m; n)) = \hat{e} & \hat{\Omega} \vdash_{\Psi; \hat{\Phi}} \hat{e} \rightsquigarrow e \Rightarrow \tau \\ \hat{\Omega} = \langle \mathcal{D}; \mathcal{G}; \mathcal{M}; \Omega_{\mathsf{app}} \rangle & \mathtt{dom}(\Omega) \cap \mathtt{dom}(\Omega_{\mathsf{app}}) = \varnothing \end{array}}{\Omega \vdash^{\hat{\Omega}; \hat{\Psi}; \hat{\Phi}; b} \mathtt{cesplicede}[m; n] \rightsquigarrow e \Rightarrow \tau} \tag{5.46m}$$

Rules (7.6a) through (7.6k) are analogous to Rules (7.1a) through (7.1k). Rule (7.6l) governs references to spliced unexpanded expressions in synthetic position, and can be understood as described in Sec. 3.2.9.

Type Analysis   Analytic ce-expression validation is governed by the following rules.

$$\frac{\Omega \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Rightarrow \tau \qquad \Omega \vdash \tau <: \tau'}{\Omega \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Leftarrow \tau'} \tag{5.47a}$$

$$\frac{\Omega \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Rightarrow \tau \qquad \Omega, x : \tau \vdash^{\mathbb{E}} \grave{e}' \rightsquigarrow e' \Leftarrow \tau'}{\Omega \vdash^{\mathbb{E}} \mathtt{celetval}(\grave{e}; x.\grave{e}') \rightsquigarrow \mathtt{ap}(\mathtt{lam}\{\tau\}(x.e'); e) \Leftarrow \tau'} \tag{5.47b}$$

$$\frac{\Gamma, x : \tau_1 \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Leftarrow \tau_2}{\Omega \vdash^{\mathbb{E}} \mathtt{ceanalam}(x.\hat{e}) \rightsquigarrow \mathtt{lam}\{\tau_1\}(x.e) \Leftarrow \mathtt{parr}(\tau_1; \tau_2)} \tag{5.47c}$$

$$\frac{\Omega \vdash^{\mathbb{C}} \grave{\kappa} \rightsquigarrow \kappa \ \mathrm{kind} \qquad \Omega, u :: \kappa \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Leftarrow \tau}{\Omega \vdash^{\mathbb{E}} \mathtt{ceclam}\{\grave{\kappa}\}(u.\grave{e}) \rightsquigarrow \mathtt{clam}\{\kappa\}(u.e) \Leftarrow \mathtt{all}\{\kappa\}(u.\tau)} \tag{5.47d}$$

$$\frac{\Omega \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Leftarrow [\mathtt{rec}(t.\tau)/t]\tau}{\Omega \vdash^{\mathbb{E}} \mathtt{cefold}(\grave{e}) \rightsquigarrow \mathtt{fold}\{t.\tau\}(e) \Leftarrow \mathtt{rec}(t.\tau)} \tag{5.47e}$$

$$\frac{\grave{e} = \mathtt{cetpl}[L](\{i \hookrightarrow \grave{e}_i\}_{i \in L}) \qquad e = \mathtt{tpl}[L](\{i \hookrightarrow e_i\}_{i \in L})}{\{\Omega \vdash^{\mathbb{E}} \grave{e}_i \rightsquigarrow e_i \Leftarrow \tau_i\}_{i \in L}}{\Omega \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Leftarrow \mathtt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L})} \tag{5.47f}$$

$$\frac{\grave{e} = \mathtt{cein}[\ell](\grave{e}') \qquad e = \mathtt{in}[L, \ell][\ell]\{\{i \hookrightarrow \tau_i\}_{i \in L}; \ell \hookrightarrow \tau\}(e')}{\Delta \, \Gamma \vdash^{\mathbb{E}} \grave{e}' \rightsquigarrow e' \Leftarrow \tau}{\Omega \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Leftarrow \mathtt{sum}[L, \ell](\{i \hookrightarrow \tau_i\}_{i \in L}; \ell \hookrightarrow \tau)} \tag{5.47g}$$

$$\frac{\Omega \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Rightarrow \tau \qquad \{\Omega \vdash^{\mathbb{E}} \grave{r}_i \rightsquigarrow r_i \Leftarrow \tau \Rightarrow \tau'\}_{1 \le i \le n}}{\Omega \vdash^{\mathbb{E}} \mathtt{cematch}[n](\grave{e}; \{\grave{r}_i\}_{1 \le i \le n}) \rightsquigarrow \mathtt{match}[n]\{\tau'\}(e; \{r_i\}_{1 \le i \le n}) \Leftarrow \tau'} \tag{5.47h}$$

$$\frac{\mathtt{parseUExp}(\mathtt{subseq}(b; m; n)) = \hat{e} \qquad \hat{\Omega} \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{e} \rightsquigarrow e \Leftarrow \tau}{\hat{\Omega} = \langle \mathcal{D}; \mathcal{G}; \mathcal{M}; \Omega_{\mathrm{app}} \rangle \qquad \mathrm{dom}(\Omega) \cap \mathrm{dom}(\Omega_{\mathrm{app}}) = \varnothing}{\Omega \vdash^{\hat{\Omega}; \hat{\Psi}; \hat{\Phi}; b} \mathtt{cesplicede}[m; n] \rightsquigarrow e \Leftarrow \tau} \tag{5.47i}$$

Rules (7.7a) through (7.7h) are analagous to Rules (7.2a) through (7.2h). Rule (7.7i) governs references to spliced unexpanded expressions in analytic position.

Bidirectional Candidate Expansion Rule Validation

The *synthetic ce-rule validation judgement* is defined mutually inductively with Rules (7.1) by the following rule.

$$\frac{\Omega \vdash p : \tau \dashv \Omega' \qquad \Omega \cup \Omega' \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Rightarrow \tau'}{\Omega \vdash^{\mathbb{E}} \mathtt{cerule}(p.\grave{e}) \rightsquigarrow \mathtt{rule}(p.e) \Rightarrow \tau \mapsto \tau'} \tag{5.48}$$

The *analytic ce-rule validation judgement* is defined mutually inductively with Rules (7.2) by the following rule.

$$\frac{\Delta \vdash p : \tau \dashv \Omega' \qquad \Omega \cup \Omega' \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Leftarrow \tau'}{\Omega \vdash^{\mathbb{E}} \mathtt{cerule}(p.\grave{e}) \rightsquigarrow \mathtt{rule}(p.e) \Leftarrow \tau \mapsto \tau'} \tag{5.49}$$

Candidate Expansion Pattern Validation

The *ce-pattern validation judgement* is inductively defined by the following rules, which are written identically to Rules (4.9).

$$\frac{}{\Omega_{\text{param}} \vdash^{\mathbb{P}} \texttt{cewildp} \rightsquigarrow \texttt{wildp} : \tau \dashv\!\!\mid^{\langle \varnothing;\varnothing;\varnothing;\varnothing \rangle}} \tag{5.50a}$$

$$\frac{\Omega_{\text{param}} \vdash^{\mathbb{P}} \grave{p} \rightsquigarrow p : [\texttt{rec}(t.\tau)/t]\tau \dashv\!\!\mid^{\hat{\Omega}}}{\Omega_{\text{param}} \vdash^{\mathbb{P}} \texttt{cefoldp}(\grave{p}) \rightsquigarrow \texttt{foldp}(p) : \texttt{rec}(t.\tau) \dashv\!\!\mid^{\hat{\Omega}}} \tag{5.50b}$$

$$\frac{\begin{array}{cc} \grave{p} = \texttt{cetplp}[L](\{i \hookrightarrow \grave{p}_i\}_{i \in L}) & p = \texttt{tplp}[L](\{i \hookrightarrow p_i\}_{i \in L}) \\ \{\Omega_{\text{param}} \vdash^{\mathbb{P}} \grave{p}_i \rightsquigarrow p_i : \tau_i \dashv\!\!\mid^{\hat{Y}_i}\}_{i \in L} \end{array}}{\Omega_{\text{param}} \vdash^{\mathbb{P}} \grave{p} \rightsquigarrow p : \texttt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L}) \dashv\!\!\mid^{\cup_{i \in L} \hat{\Omega}_i}} \tag{5.50c}$$

$$\frac{\Omega_{\text{param}} \vdash^{\mathbb{P}} \grave{p} \rightsquigarrow p : \tau \dashv\!\!\mid^{\hat{\Omega}}}{\Omega_{\text{param}} \vdash^{\mathbb{P}} \texttt{ceinp}[\ell](\grave{p}) \rightsquigarrow \texttt{inp}[\ell](p) : \texttt{sum}[L,\ell](\{i \hookrightarrow \tau_i\}_{i \in L}; \ell \hookrightarrow \tau) \dashv\!\!\mid^{\hat{\Omega}}} \tag{5.50d}$$

$$\frac{\texttt{parseUPat}(\texttt{subseq}(b;m;n)) = \hat{p} \qquad \hat{\Omega} \vdash_{\hat{\Phi}} \hat{p} \rightsquigarrow p : \tau \dashv\!\!\mid \hat{\Omega}'}{\Omega_{\text{param}} \vdash^{\hat{\Omega};\hat{\Phi};b} \texttt{cesplicedp}[m;n] \rightsquigarrow p : \tau \dashv\!\!\mid^{\hat{\Omega}'}} \tag{5.50e}$$

5.2.8   Metatheory

Chapter 6

Static Evaluation and State

In the previous sections, we have assumed that the parse functions in a TSM definition are closed expanded expressions. This is unrealistic. In this section, we discuss the semantics of the static phase of evaluation. We also add support for stateful programming with reference cells, so that we can discuss how these interact with static evaluation.

6.1   TSMs For Defining TSMs

Static functions can also make use of TSMs. In this section, we will show how quasiquotation syntax and grammar-based parser generators can be expressed using TSMs. These TSMs are quite useful for writing other TSMs.

6.1.1   Quasiquotation

TSMs must generate values of type `CEExp`. Doing so explicitly can have high syntactic cost. To decrease the syntactic cost of constructing values of this type, the prelude includes a TSM that provides quasiquotation syntax (cf. Sec. **??**):

```
syntax $qqexp at CEExp {
  static fn(body : Body) : ParseResult => (* expression parser here *)
}

syntax $qqtype at CETyp {
  static fn(body : Body) : ParseResult => (* type parser here *)
}
```

For example, the following expression:

```
let gx = $qqexp 'g(x)'
```

is more concise than its expansion:

```
let gx = App(Var 'g', Var 'x')
```

The full concrete syntax of the language can be used. Anti-quotation, i.e. splicing in an expression of type `MarkedExp`, is indicated by the prefix `%`:

```
let fgx = $qqexp 'f(%gx)'
```
The expansion of this expression is:
```
let fgx = App(Var 'f', gx)
```

6.1.2   Parser Generators

TODO: grammars, compile function, TSM for grammar, example of IP address

6.2   Static Language

We have assumed throughout this work that parse functions are fully self-contained, i.e. they are closed. This simplifies our exposition and metatheory, but it is not a realistic constraint – in practice, one would want to be able to share helper code between parse functions. To allow this, VerseML allows programmers to introduce *static blocks*, which introduce bindings available only in other static blocks and static functions. For example, the following static block defines a helper function for use in the subsequent parse function.

```
1  static
2    val parseInt : Body -> Exp option = (* ... *)
3  end
4  syntax $rx at Rx {
5    static fn(body : Body) : ParseResultExp =>
6      (* ... parseInt is available here ... *)
7  }
8  val x = parseInt("34") (* error: parseInt is not bound here *)
```

DRAFT (March 30, 2016)

Part III

TSM Implicits

Chapter 7

Unparameterized TSM Implicits

Using TSMs, a library provider can control the expansion of generalized literal forms, and thereby control the syntactic cost of common idioms. However, library clients must explicitly prefix each such form with a TSM name. In situations where the client is repeatedly using a TSM throughout a codebase, this can be inelegant. To further lower the syntactic cost of using TSMs, so that it compares to the syntactic cost of using derived forms built primitively into a language, VerseML allows clients to designate, for any type, one expression TSM and one pattern TSM as that type's *designated TSMs* within a delimited scope. When VerseML's *local type inference* system encounters a generalized literal form not prefixed by a TSM name (an *unadorned literal form*), it implicitly applies the TSM designated at the type that the expression or pattern is being checked against.

7.1 TSM Implicits By Example

We begin in this section by introducing TSM implicits by example in VerseML. In Sec. 7.2, we formalize unparameterized TSM implicits with a reduced calculus, $\mathrm{miniVerse}_{\mathsf{U}}^{\mathbf{B}}$. We will also return to the topic of TSM implicits after introducing parameterized TSMs in Chapter 5.

7.1.1 Designation

In the example in Figure 7.1, Lines 1 through 3 designate the expression TSM named `$rx`, defined in Section 3.1.2, and the pattern TSM named `$rx`, defined in Sec. 4.1.2, both at type `Rx`. These designations influence typed expansion of Lines 5 through 9.

Expression and pattern TSMs need not be designated together, nor have the same name if they are. However, this is a common idiom, so for convenience, VerseML also provides a derived designation form that combines the two designations in Figure 7.1:

```
implicit syntax $rx at Rx in (* ... *) end
```

The type annotation on a designation is technically redundant – the definition of the designated TSM determines the designated type. It is included in our examples for readability, but can be omitted if desired.

```
1  implicit syntax
2    $rx at Rx for expressions
3    $rx at Rx for patterns
4  in
5    fun is_ssn(s : string) => rx_match /\d\d\d-\d\d-\d\d\d\d/ s
6    fun name_from_example_rx(r : Rx) : string option =>
7      match r with
8        /@name: %_/ => Some name
9      | _ => None
10 end
```

Figure 7.1: An example of TSM implicits in VerseML

7.1.2 Usage

On Line 5 of Figure 7.1, we apply a function `rx_match` (not shown), which has type `Rx -> string -> MatchResult`, to an expression of unadorned literal form. During typed expansion, the expression TSM `$rx` is applied implicitly to this form to determine the expression's expansion, because `$rx` is the designated TSM at the argument type `Rx`.

Similarly, a pattern of unadorned literal form appears on Line 8. Because it appears in a syntactic position where it must match values of type `Rx`, the pattern TSM `$rx` is implicitly applied to determine its expansion.

7.1.3 Analytic and Synthetic Positions

During typed expansion of a subexpression, e', of an expresssion, e, we say that e' appears in *analytic position* if the type that e' must necessarily have can be determined based on the surrounding context, without examining e'. For example, an expression appearing as a function argument is in analytic position because the function's type determines the argument's type. Similarly, an expression may appear in analytic position due to a *type ascription*, either directly on the expression, or "further up" in the expression:

```
val ssn = /\d\d\d-\d\d-\d\d\d\d/ : Rx
val ssn : Rx = /\d\d\d-\d\d-\d\d\d\d/
fun ssn() : Rx => /\d\d\d-\d\d-\d\d\d\d/
```

If the type that e' must be assigned cannot be determined from context – i.e. e' must be examined to synthesize its type – we instead say that the expression appears in a *synthetic position*. For example, a top-level expression, or an expression appearing in a binding or function definition without a type ascription, appears in synthetic position.

Expressions of unadorned literal form can only appear in analytic position, because their type must be known to be able to determine the designated TSM that will control their expansion. For example, typed expansion of the following expression will fail because subexpressions of unadorned literal form appear in synthetic position:

```
let
  val ssn = /\d\d\d-\d\d-\d\d\d\d/ (* INVALID *)
  fun ssn() => /\d\d\d-\d\d-\d\d\d\d/ (* INVALID *)
in
  (* ... *)
end
```

Patterns can always be of unadorned literal form in VerseML, because the scrutinee of a match expression is always in synthetic position, and so the type of value that each pattern appearing within the match expression must match is always known without examining the pattern itself.

7.2 miniVerse$_\mathbf{U}^\mathbf{B}$

To formalize TSM implicits, we will now develop a reduced calculus called miniVerse$_\mathbf{U}^\mathbf{B}$, or "Bidirectional miniVerse$_\mathbf{U}$" (so named because it explicitly distinguishes type analysis from type synthesis during typed expansion, as explained below).

7.2.1 Inner Core

The inner core of miniVerse$_\mathbf{U}^\mathbf{B}$ is the same as the inner core of miniVerse$_\mathbf{U}$, as described in Sections 4.2.1 through 4.2.3. It consists of types, τ, expanded expressions, e, expanded rules, r, and expanded patterns, p.

7.2.2 Syntax of the Outer Surface

A miniVerse$_\mathbf{U}^\mathbf{B}$ program ultimately evaluates as an expanded expression. However, the programmer does not write the expanded expression directly. Instead, the programmer writes a textual sequence, b, consisting of characters in some suitable alphabet (e.g. in practice, ASCII or Unicode), which is parsed by some partial metafunction parseUExp(b) to produce an *unexpanded expression*, $\hat{e}$. Unexpanded expressions can contain *unexpanded types*, $\hat{\tau}$, *unexpanded rules*, $\hat{r}$, and *unexpanded patterns*, $\hat{p}$, so we also need partial metafunctions parseUTyp(b), parseURule(b) and parseUPat(b). The abstract syntax of unexpanded types, expressions, rules and patterns, which form the *outer surface* of miniVerse$_\mathbf{U}^\mathbf{B}$, is defined in Figure 4.3. The full definition of the textual syntax of miniVerse$_\mathbf{U}^\mathbf{B}$ is not important for our purposes, so we simply give the following condition, which states that there is some way to textually represent every unexpanded type, expression, rule and pattern.

Condition 7.1 (Textual Representability)**.** *All of the following must hold:*

1. *For each $\hat{\tau}$, there exists b such that* parseUTyp(b) $= \hat{\tau}$.
2. *For each $\hat{e}$, there exists b such that* parseUExp(b) $= \hat{e}$.
3. *For each $\hat{r}$, there exists b such that* parseURule(b) $= \hat{r}$.
4. *For each $\hat{p}$, there exists b such that* parseUPat(b) $= \hat{p}$.

Sort		Operational Form	Stylized Form	Description
UTyp	$\hat{\tau}$ ::=	$\hat{t}$	$\hat{t}$	sigil
		$\texttt{uparr}(\hat{\tau};\hat{\tau})$	$\hat{\tau} \rightharpoonup \hat{\tau}$	partial function
		$\texttt{uall}(\hat{t}.\hat{\tau})$	$\forall \hat{t}.\hat{\tau}$	polymorphic
		$\texttt{urec}(\hat{t}.\hat{\tau})$	$\mu\hat{t}.\hat{\tau}$	recursive
		$\texttt{uprod}[L](\{i \hookrightarrow \hat{\tau}_i\}_{i \in L})$	$\langle\{i \hookrightarrow \hat{\tau}_i\}_{i \in L}\rangle$	labeled product
		$\texttt{usum}[L](\{i \hookrightarrow \hat{\tau}_i\}_{i \in L})$	$[\{i \hookrightarrow \hat{\tau}_i\}_{i \in L}]$	labeled sum
UExp	$\hat{e}$ ::=	$\hat{x}$	$\hat{x}$	sigil
		$\texttt{uasc}\{\hat{\tau}\}(\hat{e})$	$\hat{e} : \hat{\tau}$	ascription
		$\texttt{uletval}(\hat{e};\hat{x}.\hat{e})$	$\texttt{let val } \hat{x} = \hat{e} \texttt{ in } \hat{e}$	value binding
		$\texttt{uanalam}(\hat{x}.\hat{e})$	$\lambda\hat{x}.\hat{e}$	abstraction (unannotated)
		$\texttt{ulam}\{\hat{\tau}\}(\hat{x}.\hat{e})$	$\lambda\hat{x}{:}\hat{\tau}.\hat{e}$	abstraction (annotated)
		$\texttt{uap}(\hat{e};\hat{e})$	$\hat{e}(\hat{e})$	application
		$\texttt{utlam}(\hat{t}.\hat{e})$	$\Lambda\hat{t}.\hat{e}$	type abstraction
		$\texttt{utap}\{\hat{\tau}\}(\hat{e})$	$\hat{e}[\hat{\tau}]$	type application
		$\texttt{ufold}(\hat{e})$	$\texttt{fold}(\hat{e})$	fold
		$\texttt{uunfold}(\hat{e})$	$\texttt{unfold}(\hat{e})$	unfold
		$\texttt{utpl}[L](\{i \hookrightarrow \hat{e}_i\}_{i \in L})$	$\langle\{i \hookrightarrow \hat{e}_i\}_{i \in L}\rangle$	labeled tuple
		$\texttt{upr}[\ell](\hat{e})$	$\hat{e} \cdot \ell$	projection
		$\texttt{uin}[\ell](\hat{e})$	$\texttt{inj } \ell \cdot \hat{e}$	injection
		$\texttt{umatch}[n](\hat{e};\{\hat{r}_i\}_{1 \leq i \leq n})$	$\texttt{match } \hat{e} \; \{\hat{r}_i\}_{1 \leq i \leq n}$	match
		$\texttt{usyntaxue}\{e\}\{\hat{\tau}\}(\hat{a}.\hat{e})$	$\texttt{syntax } \hat{a} \texttt{ at } \hat{\tau} \texttt{ for}$ $\texttt{expressions } \{e\} \texttt{ in } \hat{e}$	ueTSM definition
		$\texttt{uimplicite}[\hat{a}](\hat{e})$	$\texttt{implicit syntax } \hat{a} \texttt{ for}$ $\texttt{expressions in } \hat{e}$	ueTSM designation
		$\texttt{uapuetsm}[b][\hat{a}]$	$\hat{a} \; /b/$	ueTSM application
		$\texttt{uelit}[b]$	$/b/$	ueTSM unadorned literal
		$\texttt{usyntaxup}\{e\}\{\hat{\tau}\}(\hat{a}.\hat{e})$	$\texttt{syntax } \hat{a} \texttt{ at } \hat{\tau} \texttt{ for}$ $\texttt{patterns } \{e\} \texttt{ in } \hat{e}$	upTSM definition
		$\texttt{uimplicitp}[\hat{a}](\hat{e})$	$\texttt{implicit syntax } \hat{a} \texttt{ for}$ $\texttt{patterns in } \hat{e}$	upTSM designation
URule	$\hat{r}$ ::=	$\texttt{urule}(\hat{p}.\hat{e})$	$\hat{p} \Rightarrow \hat{e}$	match rule
UPat	$\hat{p}$ ::=	$\hat{x}$	$\hat{x}$	sigil pattern
		$\texttt{uwildp}$	$\_$	wildcard pattern
		$\texttt{ufoldp}(\hat{p})$	$\texttt{fold}(\hat{p})$	fold pattern
		$\texttt{utplp}[L](\{i \hookrightarrow \hat{p}_i\}_{i \in L})$	$\langle\{i \hookrightarrow \hat{p}_i\}_{i \in L}\rangle$	labeled tuple pattern
		$\texttt{uinp}[\ell](\hat{p})$	$\texttt{inj } \ell \cdot \hat{p}$	injection pattern
		$\texttt{uapuptsm}[b][\hat{a}]$	$\hat{a} \; /b/$	upTSM application
		$\texttt{uplit}[b]$	$/b/$	upTSM unadorned literal

Figure 7.2: Abstract syntax of unexpanded types, expressions, rules and patterns in miniVerse$_U^B$.

As in miniVerse$_U$, unexpanded types and expressions bind *type sigils*, $\hat{t}$, *expression sigils*, $\hat{x}$, and *TSM names*, $\hat{a}$. Sigils are given meaning by expansion to variables during typed expansion. We **cannot** adopt the usual definition of α-renaming of identifiers, because unexpanded types and expressions are still in a "partially parsed" state – the literal bodies, b, within an unexpanded expression might contain spliced subterms that are "surfaced" by a TSM only during typed expansion, as we will detail below.

Each inner core form (defined in Figure 4.2) maps onto an outer surface form. In particular:

- Each type variable, t, maps onto a unique type sigil, written $\hat{t}$.

- Each type form, τ, maps onto an unexpanded type form, $\mathcal{U}(\tau)$, according to the definition of $\mathcal{U}(\tau)$ in Sec. 3.2.4.

- Each expression variable, x, maps onto a unique expression sigil, written $\hat{x}$.

- Each expanded expression form, e, maps onto an unexpanded expression form $\mathcal{U}(e)$ as follows:

$$\mathcal{U}(x) = \hat{x}$$
$$\mathcal{U}(\text{lam}\{\tau\}(x.e)) = \text{ulam}\{\mathcal{U}(\tau)\}(\hat{x}.\mathcal{U}(e))$$
$$\mathcal{U}(\text{ap}(e_1; e_2)) = \text{uap}(\mathcal{U}(e_1); \mathcal{U}(e_2))$$
$$\mathcal{U}(\text{tlam}(t.e)) = \text{utlam}(\hat{t}.\mathcal{U}(e))$$
$$\mathcal{U}(\text{tap}\{\tau\}(e)) = \text{utap}\{\mathcal{U}(\tau)\}(\mathcal{U}(e))$$
$$\mathcal{U}(\text{fold}\{t.\tau\}(e)) = \text{uasc}\{\text{urec}(\hat{t}.\mathcal{U}(\tau))\}(\text{ufold}(\mathcal{U}(e)))$$
$$\mathcal{U}(\text{unfold}(e)) = \text{uunfold}(\mathcal{U}(e))$$
$$\mathcal{U}(\text{tpl}[L](\{i \hookrightarrow e_i\}_{i \in L})) = \text{utpl}[L](\{i \hookrightarrow \mathcal{U}(e_i)\}_{i \in L})$$
$$\mathcal{U}(\text{in}[L][\ell]\{\{i \hookrightarrow \tau_i\}_{i \in L}\}(e)) = \text{uasc}\{\text{usum}[L](\{i \hookrightarrow \mathcal{U}(\tau_i)\}_{i \in L})\}(\text{uin}[\ell](\mathcal{U}(e)))$$
$$\mathcal{U}(\text{match}[n]\{\tau\}(e; \{r_i\}_{1 \le i \le n})) = \text{uasc}\{\mathcal{U}(\tau)\}(\text{umatch}[n](\mathcal{U}(e); \{\mathcal{U}(r_i)\}_{1 \le i \le n}))$$

  Notice that some type arguments that appear in e appear within a type ascription in $\mathcal{U}(e)$.

- The expanded rule form maps onto the unexpanded rule form as follows:
$$\mathcal{U}(\text{rule}(p.e)) = \text{urule}(\mathcal{U}(p).\mathcal{U}(e))$$

- Each expanded pattern form, p, maps onto the unexpanded pattern form $\mathcal{U}(p)$ as follows:

$$\mathcal{U}(x) = \hat{x}$$
$$\mathcal{U}(\text{wildp}) = \text{uwildp}$$
$$\mathcal{U}(\text{foldp}(p)) = \text{ufoldp}(\mathcal{U}(p))$$
$$\mathcal{U}(\text{tplp}[L](\{i \hookrightarrow p_i\}_{i \in L})) = \text{utplp}[L](\{i \hookrightarrow \mathcal{U}(p_i)\}_{i \in L})$$
$$\mathcal{U}(\text{inp}[\ell](p)) = \text{uinp}[\ell](\mathcal{U}(p))$$

The forms related to TSM implicits are highlighted in gray in Figure 7.2.

7.2.3   Bidirectionally Typed Expansion

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

Judgement Form	Description
$\hat{\Delta} \vdash \hat{\tau} \rightsquigarrow \tau \text{ type}$	$\hat{\tau}$ is well-formed and has expansion τ assuming $\hat{\Delta}$
$\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \rightsquigarrow e \Rightarrow \tau$	$\hat{e}$ has expansion e and synthesizes type τ under $\hat{\Psi}$ and $\hat{\Phi}$ assuming $\hat{\Delta}$ and $\hat{\Gamma}$
$\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \rightsquigarrow e \Leftarrow \tau$	$\hat{e}$ has expansion e when analyzed against type τ under $\hat{\Psi}$ and $\hat{\Phi}$ assuming $\hat{\Delta}$ and $\hat{\Gamma}$
$\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{r} \rightsquigarrow r \Rightarrow \tau \Mapsto \tau'$	$\hat{r}$ has expansion r and takes values of type τ to values of synthesized type τ' under $\hat{\Psi}$ and $\hat{\Phi}$ assuming $\hat{\Delta}$ and $\hat{\Gamma}$
$\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{r} \rightsquigarrow r \Leftarrow \tau \Mapsto \tau'$	$\hat{r}$ has expansion r and takes values of type τ to values of type τ' when $\tau's$ is provided for analysis under $\hat{\Psi}$ and $\hat{\Phi}$ assuming $\hat{\Delta}$ and $\hat{\Gamma}$
$\Delta \vdash_{\hat{\Phi}} \hat{p} \rightsquigarrow p : \tau \dashv \hat{Y}$	$\hat{p}$ has expansion p and type τ and generates hypotheses $\hat{Y}$ under upTSM context $\hat{\Phi}$ assuming Δ

Type Expansion

Unexpanded type formation contexts, $\hat{\Delta}$, were defined in Sec. 3.21. The *type expansion judgement*, $\hat{\Delta} \vdash \hat{\tau} \rightsquigarrow \tau \text{ type}$, is inductively defined by Rules (3.3).

Typed Expression Expansion

In order to clearly define the semantics of TSM implicits, we must make a judgemental distinction between type synthesis and type analysis. In the latter, the type is presumed known, while in the former, it must be synthesized by examining the term that is the subject of the judgement. Expressions of unadorned literal form can only be analyzed against a known type.

The *typed expression expansion judgements*, $\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \rightsquigarrow e \Rightarrow \tau$, for type synthesis, and $\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \rightsquigarrow e \Leftarrow \tau$, for type analysis, are defined mutually inductively by Rules (7.1) and Rules (7.2), respectively, as follows.

Type Synthesis   *Unexpanded typing contexts*, $\hat{\Gamma}$, were defined in Sec. 3.2.5. Sigils that appear in $\hat{\Gamma}$ have the expansion and synthesize the type that $\hat{\Gamma}$ assigns to them.

$$\overline{\hat{\Delta}\,\hat{\Gamma}, \hat{x} \rightsquigarrow x : \tau \vdash_{\hat{\Psi};\hat{\Phi}} \hat{x} \rightsquigarrow x \Rightarrow \tau} \tag{7.1a}$$

A *type ascription* can be placed on an unexpanded expression to specify the type that it should be analyzed against. The ascribed type is synthesized if type analysis succeeds.

$$\frac{\hat{\Delta} \vdash \hat{\tau} \rightsquigarrow \tau \text{ type} \qquad \hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \rightsquigarrow e \Leftarrow \tau}{\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \mathtt{uasc}\{\hat{\tau}\}(\hat{e}) \rightsquigarrow e \Rightarrow \tau} \tag{7.1b}$$

We define let-binding of a value in synthetic position primitively in miniVerse$_\mathsf{U}^\mathbf{B}$. The following rule governs such bindings in synthetic position.

$$\frac{\hat{\Delta}\,\hat{\Gamma}\vdash_{\hat{\Psi};\hat{\Phi}}\hat{e}\rightsquigarrow e\Rightarrow\tau \qquad \hat{\Delta}\,\hat{\Gamma},\hat{x}\rightsquigarrow x:\tau\vdash_{\hat{\Psi};\hat{\Phi}}\hat{e}'\rightsquigarrow e'\Rightarrow\tau'}{\hat{\Delta}\,\hat{\Gamma}\vdash_{\hat{\Psi};\hat{\Phi}}\texttt{uletval}(\hat{e};\hat{x}.\hat{e}')\rightsquigarrow\texttt{ap}(\texttt{lam}\{\tau\}(x.e');e)\Rightarrow\tau'} \tag{7.1c}$$

Functions with an argument type annotation can appear in synthetic position.

$$\frac{\hat{\Delta}\vdash\hat{\tau}_1\rightsquigarrow\tau_1\;\texttt{type}\qquad\hat{\Delta}\,\hat{\Gamma},\hat{x}\rightsquigarrow x:\tau_1\vdash_{\hat{\Psi};\hat{\Phi}}\hat{e}\rightsquigarrow e\Rightarrow\tau_2}{\hat{\Delta}\,\hat{\Gamma}\vdash_{\hat{\Psi};\hat{\Phi}}\texttt{ulam}\{\hat{\tau}_1\}(\hat{x}.\hat{e})\rightsquigarrow\texttt{lam}\{\tau_1\}(x.e)\Rightarrow\texttt{parr}(\tau_1;\tau_2)} \tag{7.1d}$$

Function applications can appear in synthetic position. The argument is analyzed against the argument type synthesized by the function.

$$\frac{\hat{\Delta}\,\hat{\Gamma}\vdash_{\hat{\Psi};\hat{\Phi}}\hat{e}_1\rightsquigarrow e_1\Rightarrow\texttt{parr}(\tau_2;\tau)\qquad\hat{\Delta}\,\hat{\Gamma}\vdash_{\hat{\Psi};\hat{\Phi}}\hat{e}_2\rightsquigarrow e_2\Leftarrow\tau_2}{\hat{\Delta}\,\hat{\Gamma}\vdash_{\hat{\Psi};\hat{\Phi}}\texttt{uap}(\hat{e}_1;\hat{e}_2)\rightsquigarrow\texttt{ap}(e_1;e_2)\Rightarrow\tau} \tag{7.1e}$$

Type lambdas and type applications can appear in synthetic position.

$$\frac{\hat{\Delta},\hat{t}\rightsquigarrow t\;\texttt{type}\;\hat{\Gamma}\vdash_{\hat{\Psi};\hat{\Phi}}\hat{e}\rightsquigarrow e\Rightarrow\tau}{\hat{\Delta}\,\hat{\Gamma}\vdash_{\hat{\Psi};\hat{\Phi}}\texttt{utlam}(\hat{t}.\hat{e})\rightsquigarrow\texttt{tlam}(t.e)\Rightarrow\texttt{all}(t.\tau)} \tag{7.1f}$$

$$\frac{\hat{\Delta}\,\hat{\Gamma}\vdash_{\hat{\Psi};\hat{\Phi}}\hat{e}\rightsquigarrow e\Rightarrow\texttt{all}(t.\tau)\qquad\hat{\Delta}\vdash\hat{\tau}'\rightsquigarrow\tau'\;\texttt{type}}{\hat{\Delta}\,\hat{\Gamma}\vdash_{\hat{\Psi};\hat{\Phi}}\texttt{utap}\{\hat{\tau}'\}(\hat{e})\rightsquigarrow\texttt{tap}\{\tau'\}(e)\Rightarrow[\tau'/t]\tau} \tag{7.1g}$$

Unfoldings can appear in synthetic position.

$$\frac{\hat{\Delta}\,\hat{\Gamma}\vdash_{\hat{\Psi};\hat{\Phi}}\hat{e}\rightsquigarrow e\Rightarrow\texttt{rec}(t.\tau)}{\hat{\Delta}\,\hat{\Gamma}\vdash_{\hat{\Psi};\hat{\Phi}}\texttt{uunfold}(\hat{e})\rightsquigarrow\texttt{unfold}(e)\Rightarrow[\texttt{rec}(t.\tau)/t]\tau} \tag{7.1h}$$

Labeled tuples can appear in synthetic position. Each of the field values are then in synthetic position.

$$\frac{\{\hat{\Delta}\,\hat{\Gamma}\vdash_{\hat{\Psi};\hat{\Phi}}\hat{e}_i\rightsquigarrow e_i\Rightarrow\tau_i\}_{i\in L}}{\hat{\Delta}\,\hat{\Gamma}\vdash_{\hat{\Psi};\hat{\Phi}}\texttt{utpl}[L](\{i\hookrightarrow\hat{e}_i\}_{i\in L})\rightsquigarrow\texttt{tpl}[L](\{i\hookrightarrow e_i\}_{i\in L})\Rightarrow\texttt{prod}[L](\{i\hookrightarrow\tau_i\}_{i\in L})} \tag{7.1i}$$

Fields can be projected out of a labeled tuple in synthetic position.

$$\frac{\hat{\Delta}\,\hat{\Gamma}\vdash_{\hat{\Psi};\hat{\Phi}}\hat{e}\rightsquigarrow e\Rightarrow\texttt{prod}[L,\ell](\{i\hookrightarrow\tau_i\}_{i\in L};\ell\hookrightarrow\tau)}{\hat{\Delta}\,\hat{\Gamma}\vdash_{\hat{\Psi};\hat{\Phi}}\texttt{upr}[\ell](\hat{e})\rightsquigarrow\texttt{pr}[\ell](e)\Rightarrow\tau} \tag{7.1j}$$

Match expressions can appear in synthetic position, as long as there is at least one rule.

$$\frac{n>0\qquad\hat{\Delta}\,\hat{\Gamma}\vdash_{\hat{\Psi};\hat{\Phi}}\hat{e}\rightsquigarrow e\Rightarrow\tau\qquad\{\hat{\Delta}\,\hat{\Gamma}\vdash_{\hat{\Psi};\hat{\Phi}}\hat{r}_i\rightsquigarrow r_i\Rightarrow\tau\Mapsto\tau'\}_{1\leq i\leq n}}{\hat{\Delta}\,\hat{\Gamma}\vdash_{\hat{\Psi};\hat{\Phi}}\texttt{umatch}[n](\hat{e};\{\hat{r}_i\}_{1\leq i\leq n})\rightsquigarrow\texttt{match}[n]\{\tau'\}(e;\{r_i\}_{1\leq i\leq n})\Rightarrow\tau'} \tag{7.1k}$$

ueTSMs can be defined and applied in synthetic position.

$$\frac{\hat{\Delta} \vdash \hat{\tau} \rightsquigarrow \tau \, \mathtt{type} \qquad \emptyset \, \emptyset \vdash e_{\mathrm{parse}} : \mathtt{parr(Body;ParseResultExp)}}{\hat{\Delta} \, \hat{\Gamma} \vdash_{\hat{\Psi}, \hat{a} \rightsquigarrow a \hookrightarrow \mathtt{uetsm}(\tau; e_{\mathrm{parse}}); \hat{\Phi}} \, \hat{e} \rightsquigarrow e \Rightarrow \tau'} \tag{7.1l}$$

$$\hat{\Delta} \, \hat{\Gamma} \vdash_{\hat{\Psi}; \hat{\Phi}} \mathtt{syntax} \, \hat{a} \, \mathtt{at} \, \hat{\tau} \, \mathtt{for} \, \mathtt{expressions} \, \{e_{\mathrm{parse}}\} \, \mathtt{in} \, \hat{e} \rightsquigarrow e \Rightarrow \tau'$$

$$\frac{b \downarrow e_{\mathrm{body}} \qquad e_{\mathrm{parse}}(e_{\mathrm{body}}) \Downarrow \mathtt{inj} \, \mathtt{Success} \cdot e_{\mathrm{cand}} \qquad e_{\mathrm{cand}} \uparrow_{\mathsf{CEExp}} \grave{e}}{\emptyset \, \emptyset \vdash^{\hat{\Delta}; \hat{\Gamma}; \hat{\Psi}, \hat{a} \rightsquigarrow a \hookrightarrow \mathtt{uetsm}(\tau; e_{\mathrm{parse}}); \hat{\Phi}; b} \, \grave{e} \rightsquigarrow e \Leftarrow \tau}{\hat{\Delta} \, \hat{\Gamma} \vdash_{\hat{\Psi}, \hat{a} \rightsquigarrow a \hookrightarrow \mathtt{uetsm}(\tau; e_{\mathrm{parse}}); \hat{\Phi}} \, \hat{a} \, /b/ \rightsquigarrow e \Rightarrow \tau} \tag{7.1m}$$

These rules are nearly identical to Rules (4.4l) and (4.4m), differing only in that the typed expansion premises have been replaced by corresponding synthetic typed expansion premises. The premises of these rules can be understood as described in Sections 3.2.6 and 3.2.7. The body encoding judgement and candidate expansion expression decoding judgements were characterized in Sec. 4.2.5. We discuss candidate expansion validation in Sec. 7.2.5 below.

To support ueTSM implicits, ueTSM contexts, $\hat{\Psi}$, are redefined to take the form $\langle \mathcal{A}; \Psi; \mathcal{I} \rangle$. TSM naming contexts, $\mathcal{A}$, and ueTSM definition contexts, Ψ, were defined in Sec. 4.2.5. We write $\hat{\Psi}, \hat{a} \rightsquigarrow a \hookrightarrow \mathtt{uetsm}(\tau; e_{\mathrm{parse}})$ when $\hat{\Psi} = \langle \mathcal{A}; \Psi; \mathcal{I} \rangle$ as shorthand for

$$\langle \mathcal{A} \uplus \hat{a} \rightsquigarrow a; \Psi, a \hookrightarrow \mathtt{uetsm}(\tau; e_{\mathrm{parse}}); \mathcal{I} \rangle$$

TSM designation contexts, $\mathcal{I}$, are finite functions that map each type $\tau \in \mathrm{dom}(\mathcal{I})$ to the *TSM designation* $\tau \hookrightarrow a$, for some symbol a. We write $\mathcal{I} \uplus \tau \hookrightarrow a$ for the TSM designation context that maps τ to $\tau \hookrightarrow a$ and defers to $\mathcal{I}$ for all other types (i.e. the previous designation, if any, is updated).

The TSM designation context in the ueTSM context is updated by expressions of ueTSM designation form. Such expressions can appear in synthetic position, where they are governed by the following rule:

$$\frac{\hat{\Delta} \, \hat{\Gamma} \vdash_{\langle \mathcal{A} \uplus \hat{a} \rightsquigarrow a; \Psi, a \hookrightarrow \mathtt{uetsm}(\tau; e_{\mathrm{parse}}); \mathcal{I} \uplus \tau \hookrightarrow a \rangle; \hat{\Phi}} \, \hat{e} \rightsquigarrow e \Rightarrow \tau'}{\hat{\Delta} \, \hat{\Gamma} \vdash_{\langle \mathcal{A} \uplus \hat{a} \rightsquigarrow a; \Psi, a \hookrightarrow \mathtt{uetsm}(\tau; e_{\mathrm{parse}}); \mathcal{I} \rangle; \hat{\Phi}} \, \mathtt{implicit} \, \mathtt{syntax} \, \hat{a} \, \mathtt{for} \, \mathtt{expressions} \, \mathtt{in} \, \hat{e} \rightsquigarrow e \Rightarrow \tau'} \tag{7.1n}$$

Like ueTSMs, upTSMs can be defined in synthetic position.

$$\frac{\hat{\Delta} \vdash \hat{\tau} \rightsquigarrow \tau \, \mathtt{type} \qquad \emptyset \, \emptyset \vdash e_{\mathrm{parse}} : \mathtt{parr(Body;ParseResultPat)}}{\hat{\Delta} \, \hat{\Gamma} \vdash_{\hat{\Psi}; \hat{\Phi}, \hat{a} \rightsquigarrow a \hookrightarrow \mathtt{uptsm}(\tau; e_{\mathrm{parse}})} \, \hat{e} \rightsquigarrow e \Rightarrow \tau'}{\hat{\Delta} \, \hat{\Gamma} \vdash_{\hat{\Psi}; \hat{\Phi}} \mathtt{syntax} \, \hat{a} \, \mathtt{at} \, \hat{\tau} \, \mathtt{for} \, \mathtt{patterns} \, \{e_{\mathrm{parse}}\} \, \mathtt{in} \, \hat{e} \rightsquigarrow e \Rightarrow \tau'} \tag{7.1o}$$

This rule is nearly identical to Rule (4.4n), differing only in that the typed expansion premise has been replaced by the corresponding synthetic typed expansion premise. The premises can be understood as described in Section 4.2.6.

To support upTSM implicits, upTSM contexts, $\hat{\Phi}$, are redefined to take the form $\langle \mathcal{A}; \Phi; \mathcal{I} \rangle$. upTSM definition contexts, Φ, were defined in Sec. 4.2.6. We write $\hat{\Phi}, \hat{a} \rightsquigarrow a \hookrightarrow \mathtt{uptsm}(\tau; e_{\mathrm{parse}})$ when $\hat{\Phi} = \langle \mathcal{A}; \Phi; \mathcal{I} \rangle$ as shorthand for

$$\langle \mathcal{A} \uplus \hat{a} \rightsquigarrow a; \Phi, a \hookrightarrow \mathtt{uptsm}(\tau; e_{\mathrm{parse}}); \mathcal{I} \rangle$$

The TSM designation context in the upTSM context is updated by expressions of upTSM designation form. Such expressions can appear in synthetic position, where they are governed by the following rule:

$$\frac{\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi}; \langle \mathcal{A} \uplus \hat{a} \rightsquigarrow a; \Phi, a \hookrightarrow \mathtt{uptsm}(\tau; e_{\mathrm{parse}}); \mathcal{I} \uplus \tau \hookrightarrow a \rangle} \hat{e} \rightsquigarrow e \Rightarrow \tau'}{\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi}; \langle \mathcal{A} \uplus \hat{a} \rightsquigarrow a; \Phi, a \hookrightarrow \mathtt{uetsm}(\tau; e_{\mathrm{parse}}); \mathcal{I} \rangle} \mathtt{implicit\ syntax}\ \hat{a}\ \mathtt{for\ patterns\ in}\ \hat{e} \rightsquigarrow e \Rightarrow \tau'} \tag{7.1p}$$

Type Analysis Type analysis subsumes type synthesis, in that when a type can be synthesized for an unexpanded expression, that unexpanded expression can also be analyzed against that type, producing the same expansion. This is expressed by the following *subsumption rule* for unexpanded expressions.

$$\frac{\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{e} \rightsquigarrow e \Rightarrow \tau}{\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{e} \rightsquigarrow e \Leftarrow \tau} \tag{7.2a}$$

Additional rules are needed for certain forms in order to propagate types for analysis into subexpressions, and for forms that can appear only in analytic position.

Rule (7.1c) governed value bindings in synthetic position. The following rule governs value bindings in analytic position.

$$\frac{\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{e} \rightsquigarrow e \Rightarrow \tau \qquad \hat{\Delta}\,\hat{\Gamma}, \hat{x} \rightsquigarrow x : \tau \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{e}' \rightsquigarrow e' \Leftarrow \tau'}{\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi}; \hat{\Phi}} \mathtt{uletval}(\hat{e}; \hat{x}.\hat{e}') \rightsquigarrow \mathtt{ap}(\mathtt{lam}\{\tau\}(x.e'); e) \Leftarrow \tau'} \tag{7.2b}$$

An unannotated function can appear only in analytic position. The argument type is determined from the type that the unannotated function is being analyzed against.

$$\frac{\hat{\Delta}\,\hat{\Gamma}, \hat{x} \rightsquigarrow x : \tau_1 \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{e} \rightsquigarrow e \Leftarrow \tau_2}{\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi}; \hat{\Phi}} \mathtt{uanalam}(\hat{x}.\hat{e}) \rightsquigarrow \mathtt{lam}\{\tau_1\}(x.e) \Leftarrow \mathtt{parr}(\tau_1; \tau_2)} \tag{7.2c}$$

Rule (7.1f) governed type lambdas in synthetic position. The following rule governs type lambdas in analytic position.

$$\frac{\hat{\Delta}, \hat{t} \rightsquigarrow t\ \mathtt{type}\ \hat{\Gamma} \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{e} \rightsquigarrow e \Leftarrow \tau}{\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi}; \hat{\Phi}} \mathtt{utlam}(\hat{t}.\hat{e}) \rightsquigarrow \mathtt{tlam}(t.e) \Leftarrow \mathtt{all}(t.\tau)} \tag{7.2d}$$

Values of recursive types can be introduced only in analytic position.

$$\frac{\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{e} \rightsquigarrow e \Leftarrow [\mathtt{rec}(t.\tau)/t]\tau}{\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi}; \hat{\Phi}} \mathtt{ufold}(\hat{e}) \rightsquigarrow \mathtt{fold}\{t.\tau\}(e) \Leftarrow \mathtt{rec}(t.\tau)} \tag{7.2e}$$

Rule (7.1i) governed labeled tuples in synthetic position. The following rule governs labeled tuples in analytic position.

$$\frac{\{\hat{\Delta}\,\hat{\Gamma}\vdash_{\Psi;\hat{\Phi}}\hat{e}_i\rightsquigarrow e_i\Leftarrow\tau_i\}_{i\in L}}{\hat{\Delta}\,\hat{\Gamma}\vdash_{\Psi;\hat{\Phi}}\mathtt{utpl}[L](\{i\hookrightarrow\hat{e}_i\}_{i\in L})\rightsquigarrow\mathtt{tpl}[L](\{i\hookrightarrow e_i\}_{i\in L})\Leftarrow\mathtt{prod}[L](\{i\hookrightarrow\tau_i\}_{i\in L})}$$
(7.2f)

Values of labeled sum type can appear only in analytic position.

$$\left(\frac{\hat{\Delta}\,\hat{\Gamma}\vdash_{\Psi;\hat{\Phi}}\hat{e}\rightsquigarrow e\Leftarrow\tau}{\underset{\rightsquigarrow}{\hat{\Delta}\,\hat{\Gamma}\vdash_{\Psi;\hat{\Phi}}\mathtt{uin}[\ell](\hat{e})}\atop\mathtt{in}[L,\ell][\ell]\{\{i\hookrightarrow\tau_i\}_{i\in L};\ell\hookrightarrow\tau\}(e)\Leftarrow\mathtt{sum}[L,\ell](\{i\hookrightarrow\tau_i\}_{i\in L};\ell\hookrightarrow\tau)}\right)$$
(7.2g)

Rule (7.1k) governed match expressions in synthetic position. The following rule governs match expressions in analytic position.

$$\frac{\hat{\Delta}\,\hat{\Gamma}\vdash_{\Psi;\hat{\Phi}}\hat{e}\rightsquigarrow e\Rightarrow\tau\qquad\{\hat{\Delta}\,\hat{\Gamma}\vdash_{\Psi;\hat{\Phi}}\hat{r}_i\rightsquigarrow r_i\Leftarrow\tau\mapsto\tau'\}_{1\leq i\leq n}}{\hat{\Delta}\,\hat{\Gamma}\vdash_{\Psi;\hat{\Phi}}\mathtt{umatch}[n](\hat{e};\{\hat{r}_i\}_{1\leq i\leq n})\rightsquigarrow\mathtt{match}[n]\{\tau'\}(e;\{r_i\}_{1\leq i\leq n})\Leftarrow\tau'}$$
(7.2h)

Rule (7.1l) governed ueTSM definitions in synthetic position. The following rule governs ueTSM definitions in analytic position.

$$\frac{\hat{\Delta}\vdash\hat{\tau}\rightsquigarrow\tau\ \mathrm{type}\qquad\varnothing\,\varnothing\vdash e_{\mathrm{parse}}:\mathtt{parr}(\mathtt{Body};\mathtt{ParseResultExp})}{\hat{\Delta}\,\hat{\Gamma}\vdash_{\Psi,\hat{a}\rightsquigarrow a\hookrightarrow\mathtt{uetsm}(\tau;e_{\mathrm{parse}});\hat{\Phi}}\hat{e}\rightsquigarrow e\Leftarrow\tau'}{\hat{\Delta}\,\hat{\Gamma}\vdash_{\Psi;\hat{\Phi}}\mathtt{syntax}\ \hat{a}\ \mathtt{at}\ \hat{\tau}\ \mathtt{for}\ \mathtt{expressions}\ \{e_{\mathrm{parse}}\}\ \mathtt{in}\ \hat{e}\rightsquigarrow e\Leftarrow\tau'}$$
(7.2i)

Rule (7.1n) governed ueTSM designations in synthetic position. The following rule governs ueTSM designations in analytic position.

$$\frac{\hat{\Delta}\,\hat{\Gamma}\vdash_{\langle\mathcal{A}\uplus\hat{a}\rightsquigarrow a;\Psi,a\hookrightarrow\mathtt{uetsm}(\tau;e_{\mathrm{parse}});\mathcal{I}\uplus\tau\hookrightarrow a\rangle;\hat{\Phi}}\hat{e}\rightsquigarrow e\Leftarrow\tau'}{\hat{\Delta}\,\hat{\Gamma}\vdash_{\langle\mathcal{A}\uplus\hat{a}\rightsquigarrow a;\Psi,a\hookrightarrow\mathtt{uetsm}(\tau;e_{\mathrm{parse}});\mathcal{I}\rangle;\hat{\Phi}}\mathtt{implicit}\ \mathtt{syntax}\ \hat{a}\ \mathtt{for}\ \mathtt{expressions}\ \mathtt{in}\ \hat{e}\rightsquigarrow e\Leftarrow\tau'}$$
(7.2j)

An expression of unadorned literal form can appear only in analytic position. The following rule extracts the TSM designated at the type that the expression is being analyzed against from the TSM designation context in the ueTSM context and applies it implicitly, i.e. the premises correspond to those of Rule (7.1m).

$$\frac{b\downarrow e_{\mathrm{body}}\qquad e_{\mathrm{parse}}(e_{\mathrm{body}})\Downarrow\mathtt{inj}\ \mathtt{Success}\cdot e_{\mathrm{cand}}\qquad e_{\mathrm{cand}}\uparrow_{\mathrm{CEExp}}\grave{e}}{\varnothing\,\varnothing\vdash^{\hat{\Delta};\hat{\Gamma};\langle\mathcal{A};\Psi,a\hookrightarrow\mathtt{uetsm}(\tau;e_{\mathrm{parse}});\mathcal{I}\uplus\tau\hookrightarrow a\rangle;\hat{\Phi};b}\grave{e}\rightsquigarrow e\Leftarrow\tau}{\hat{\Delta}\,\hat{\Gamma}\vdash_{\langle\mathcal{A};\Psi,a\hookrightarrow\mathtt{uetsm}(\tau;e_{\mathrm{parse}});\mathcal{I}\uplus\tau\hookrightarrow a\rangle;\hat{\Phi}}\mathtt{uelit}[b]\rightsquigarrow e\Leftarrow\tau}$$
(7.2k)

Rule (7.1o) governed upTSM definitions in synthetic position. The following rule governs upTSM definitions in analytic position.

$$\frac{\hat{\Delta}\vdash\hat{\tau}\rightsquigarrow\tau\ \mathrm{type}\qquad\varnothing\,\varnothing\vdash e_{\mathrm{parse}}:\mathtt{parr}(\mathtt{Body};\mathtt{ParseResultPat})}{\hat{\Delta}\,\hat{\Gamma}\vdash_{\Psi;\hat{\Phi},\hat{a}\rightsquigarrow a\hookrightarrow\mathtt{uptsm}(\tau;e_{\mathrm{parse}})}\hat{e}\rightsquigarrow e\Leftarrow\tau'}{\hat{\Delta}\,\hat{\Gamma}\vdash_{\Psi;\hat{\Phi}}\mathtt{syntax}\ \hat{a}\ \mathtt{at}\ \hat{\tau}\ \mathtt{for}\ \mathtt{patterns}\ \{e_{\mathrm{parse}}\}\ \mathtt{in}\ \hat{e}\rightsquigarrow e\Leftarrow\tau'}$$
(7.2l)

Rule (7.1p) governed upTSM designations in synthetic position. The following rule governs upTSM designations in analytic position.

$$\frac{\hat{\Delta}\,\hat{\Gamma}\vdash_{\hat{\Psi};\langle\mathcal{A}\uplus\hat{a}\rightsquigarrow a;\Phi,a\hookrightarrow\mathtt{uptsm}(\tau;e_{\mathrm{parse}});\mathcal{I}\uplus\tau\hookrightarrow a\rangle}\hat{e}\rightsquigarrow e\Leftarrow\tau'}{\hat{\Delta}\,\hat{\Gamma}\vdash_{\hat{\Psi};\langle\mathcal{A}\uplus\hat{a}\rightsquigarrow a;\Phi,a\hookrightarrow\mathtt{uetsm}(\tau;e_{\mathrm{parse}});\mathcal{I}\rangle}\ \mathtt{implicit\ syntax}\ \hat{a}\ \mathtt{for\ patterns\ in}\ \hat{e}\rightsquigarrow e\Leftarrow\tau'}$$

$$(7.2\mathrm{m})$$

Typed Rule Expansion

The synthetic typed rule expansion judgement is invoked iteratively by Rule (7.1k) to synthesize a type, τ', from the branch expressions in the rule sequence. This judgement is defined mutually inductively with Rules (7.1) and Rules (7.2) by the following rule.

$$\frac{\Delta\vdash_{\hat{\Phi}}\hat{p}\rightsquigarrow p:\tau\dashv\langle\mathcal{G}';\Gamma'\rangle\qquad\langle\mathcal{D};\Delta\rangle\,\langle\mathcal{G}\uplus\mathcal{G}';\Gamma\cup\Gamma'\rangle\vdash_{\hat{\Psi};\hat{\Phi}}\hat{e}\rightsquigarrow e\Rightarrow\tau'}{\langle\mathcal{D};\Delta\rangle\,\langle\mathcal{G};\Gamma\rangle\vdash_{\hat{\Psi};\hat{\Phi}}\mathtt{urule}(\hat{p}.\hat{e})\rightsquigarrow\mathtt{rule}(p.e)\Rightarrow\tau\mapsto\tau'}\qquad(7.3)$$

The analytic typed rule expansion judgement is invoked iteratively by Rule (7.2h). This judgement is defined mutually inductively with Rules (7.1), Rules (7.2), and Rule (7.3) by the following rule, which is the analytic analag of Rule (7.3).

$$\frac{\Delta\vdash_{\hat{\Phi}}\hat{p}\rightsquigarrow p:\tau\dashv\langle\mathcal{G}';\Gamma'\rangle\qquad\langle\mathcal{D};\Delta\rangle\,\langle\mathcal{G}\uplus\mathcal{G}';\Gamma\cup\Gamma'\rangle\vdash_{\hat{\Psi};\hat{\Phi}}\hat{e}\rightsquigarrow e\Leftarrow\tau'}{\langle\mathcal{D};\Delta\rangle\,\langle\mathcal{G};\Gamma\rangle\vdash_{\hat{\Psi};\hat{\Phi}}\mathtt{urule}(\hat{p}.\hat{e})\rightsquigarrow\mathtt{rule}(p.e)\Leftarrow\tau\mapsto\tau'}\qquad(7.4)$$

The premises of these rules can be understood as described in Sec. 4.2.5.

Typed Pattern Expansion

The typed pattern expansion judgement is inductively defined by Rules (7.5) as follows.
   The following rules are written identically to the typed pattern expansion rules for shared pattern forms in miniVerse$_\mathsf{U}$, i.e. Rules (4.6a) through (4.6e).

$$\frac{}{\Delta\vdash_{\hat{\Phi}}\hat{x}\rightsquigarrow x:\tau\dashv\langle\hat{x}\rightsquigarrow x;x:\tau\rangle}\qquad(7.5\mathrm{a})$$

$$\frac{}{\Delta\vdash_{\hat{\Phi}}\mathtt{uwildp}\rightsquigarrow\mathtt{wildp}:\tau\dashv\langle\varnothing;\varnothing\rangle}\qquad(7.5\mathrm{b})$$

$$\frac{\Delta\vdash_{\hat{\Phi}}\hat{p}\rightsquigarrow p:[\mathtt{rec}(t.\tau)/t]\tau\dashv\hat{Y}}{\Delta\vdash_{\hat{\Phi}}\mathtt{ufoldp}(\hat{p})\rightsquigarrow\mathtt{foldp}(p):\mathtt{rec}(t.\tau)\dashv\hat{Y}}\qquad(7.5\mathrm{c})$$

$$\frac{\{\Delta\vdash_{\hat{\Phi}}\hat{p}_i\rightsquigarrow p_i:\tau_i\dashv\hat{Y}_i\}_{i\in L}}{\left(\begin{array}{c}\Delta\vdash_{\hat{\Phi}}\underset{\rightsquigarrow}{\mathtt{utplp}[L](\{i\hookrightarrow\hat{p}_i\}_{i\in L})}\\\mathtt{tplp}[L](\{i\hookrightarrow p_i\}_{i\in L}):\mathtt{prod}[L](\{i\hookrightarrow\tau_i\}_{i\in L})\dashv\cup_{i\in L}\hat{Y}_i\end{array}\right)}\qquad(7.5\mathrm{d})$$

$$\frac{\Delta\vdash_{\hat{\Phi}}\hat{p}\rightsquigarrow p:\tau\dashv\hat{Y}}{\Delta\vdash_{\hat{\Phi}}\mathtt{uinp}[\ell](\hat{p})\rightsquigarrow\mathtt{inp}[\ell](p):\mathtt{sum}[L,\ell](\{i\hookrightarrow\tau_i\}_{i\in L};\ell\hookrightarrow\tau)\dashv\hat{Y}}\qquad(7.5\mathrm{e})$$

The following rule governs upTSM application. It is written identically to Rule (4.6f).

$$\frac{b \downarrow e_{\text{body}} \qquad e_{\text{parse}}(e_{\text{body}}) \Downarrow \texttt{inj Success} \cdot e_{\text{cand}} \qquad e_{\text{cand}} \uparrow_{\text{CEPat}} \grave{p} \qquad \vdash^{\Delta; \hat{\Phi}, \hat{a} \rightsquigarrow a \hookrightarrow \texttt{uptsm}(\tau; e_{\text{parse}}); b}\ \grave{p} \rightsquigarrow p : \tau \dashv\!\!\mid^{\hat{Y}}}{\Delta \vdash_{\hat{\Phi}, \hat{a} \rightsquigarrow a \hookrightarrow \texttt{uptsm}(\tau; e_{\text{parse}})} \hat{a} \ /b/ \rightsquigarrow p : \tau \dashv\!\!\mid \hat{Y}} \tag{7.5f}$$

Unexpanded patterns of unadorned literal form are governed by the following rule, which extracts the designated upTSM from the upTSM context and applies it implicitly, i.e. the premises correspond to those of Rule (7.5f).

$$\frac{b \downarrow e_{\text{body}} \qquad e_{\text{parse}}(e_{\text{body}}) \Downarrow \texttt{inj Success} \cdot e_{\text{cand}} \qquad e_{\text{cand}} \uparrow_{\text{CEPat}} \grave{p} \qquad \vdash^{\Delta; \langle \mathcal{A}; \Phi, a \hookrightarrow \texttt{uptsm}(\tau; e_{\text{parse}}); \mathcal{I}, \tau \hookrightarrow a \rangle; b}\ \grave{p} \rightsquigarrow p : \tau \dashv\!\!\mid^{\hat{Y}}}{\Delta \vdash_{\langle \mathcal{A}; \Phi, a \hookrightarrow \texttt{uptsm}(\tau; e_{\text{parse}}); \mathcal{I}, \tau \hookrightarrow a \rangle} /b/ \rightsquigarrow p : \tau \dashv\!\!\mid \hat{Y}} \tag{7.5g}$$

Outer Surface Expressibility

The following lemma establishes that each well-typed expanded pattern can be expressed as an unexpanded pattern matching values of the same type and generating the same hypotheses and corresponding sigil updates. The metafunction $\mathcal{U}(Y)$ was defined in 4.2.5.

Lemma 7.2 (Pattern Expressibility). *If* $\Delta \vdash p : \tau \dashv\!\!\mid Y$ *then* $\Delta \vdash_{\hat{\Phi}} \mathcal{U}(p) \rightsquigarrow p : \tau \dashv\!\!\mid \mathcal{U}(Y)$.

Proof. By rule induction over Rules (4.3), using the definitions of $\mathcal{U}(Y)$ and $\mathcal{U}(p)$. In each case, we can apply the IH to or over each premise, then apply the corresponding rule in Rules (7.5). $\square$

We can now establish the Expressibility Theorem – that each well-typed expanded expression, e, can be expressed as an unexpanded expression, $\hat{e}$, which synthesizes the same type under the corresponding contexts.

Theorem 7.3 (Expressibility). *Both of the following hold:*

1. *If* $\Delta \, \Gamma \vdash e : \tau$ *then* $\mathcal{U}(\Delta) \, \mathcal{U}(\Gamma) \vdash_{\hat{\Psi}; \hat{\Phi}} \mathcal{U}(e) \rightsquigarrow e \Rightarrow \tau$.
2. *If* $\Delta \, \Gamma \vdash r : \tau \Mapsto \tau'$ *then* $\mathcal{U}(\Delta) \, \mathcal{U}(\Gamma) \vdash_{\hat{\Psi}; \hat{\Phi}} \mathcal{U}(r) \rightsquigarrow r \Rightarrow \tau \Mapsto \tau'$.

Proof. By mutual rule induction over Rules (4.1) and Rule (4.2) using the definitions of $\mathcal{U}(\Delta)$, $\mathcal{U}(\Gamma)$, $\mathcal{U}(e)$ and $\mathcal{U}(r)$. In each case, we apply the IH, part 1 to or over each typing premise, the IH, part 2 over each rule typing premise, Lemma 3.13 to or over each type formation premise, Lemma 7.2 to each pattern typing premise, then derive the conclusion by applying Rules (7.1) and Rule (7.3). $\square$

Sort		Operational Form	Stylized Form	Description
CETyp	$\tilde{\tau} ::=$	t	t	variable
		$\mathtt{ceparr}(\tilde{\tau};\tilde{\tau})$	$\tilde{\tau} \rightharpoonup \tilde{\tau}$	partial function
		$\mathtt{ceall}(t.\tilde{\tau})$	$\forall t.\tilde{\tau}$	polymorphic
		$\mathtt{cerec}(t.\tilde{\tau})$	$\mu t.\tilde{\tau}$	recursive
		$\mathtt{ceprod}[L](\{i \hookrightarrow \tilde{\tau}_i\}_{i \in L})$	$\langle \{i \hookrightarrow \tilde{\tau}_i\}_{i \in L} \rangle$	labeled product
		$\mathtt{cesum}[L](\{i \hookrightarrow \tilde{\tau}_i\}_{i \in L})$	$[\{i \hookrightarrow \tilde{\tau}_i\}_{i \in L}]$	labeled sum
		$\mathtt{cesplicedt}[m;n]$	$\mathtt{splicedt}\langle m, n \rangle$	spliced
CEExp	$\grave{e} ::=$	x	x	variable
		$\mathtt{ceasc}\{\tilde{\tau}\}(\grave{e})$	$\grave{e} : \tilde{\tau}$	ascription
		$\mathtt{celetval}(\grave{e};x.\grave{e})$	$\mathtt{let\ val}\ x = \grave{e}\ \mathtt{in}\ \grave{e}$	value binding
		$\mathtt{ceanalam}(x.\grave{e})$	$\lambda x.\grave{e}$	abstraction (unannotated)
		$\mathtt{celam}\{\tilde{\tau}\}(x.\grave{e})$	$\lambda x{:}\tilde{\tau}.\grave{e}$	abstraction (annotated)
		$\mathtt{ceap}(\grave{e};\grave{e})$	$\grave{e}(\grave{e})$	application
		$\mathtt{cetlam}(t.\grave{e})$	$\Lambda t.\grave{e}$	type abstraction
		$\mathtt{cetap}\{\tilde{\tau}\}(\grave{e})$	$\grave{e}[\tilde{\tau}]$	type application
		$\mathtt{cefold}(\grave{e})$	$\mathtt{fold}(\grave{e})$	fold
		$\mathtt{ceunfold}(\grave{e})$	$\mathtt{unfold}(\grave{e})$	unfold
		$\mathtt{cetpl}[L](\{i \hookrightarrow \grave{e}_i\}_{i \in L})$	$\langle \{i \hookrightarrow \grave{e}_i\}_{i \in L} \rangle$	labeled tuple
		$\mathtt{cepr}[\ell](\grave{e})$	$\grave{e} \cdot \ell$	projection
		$\mathtt{cein}[\ell](\grave{e})$	$\mathtt{inj}\ \ell \cdot \grave{e}$	injection
		$\mathtt{cematch}[n](\grave{e};\{\grave{r}_i\}_{1 \leq i \leq n})$	$\mathtt{match}\ \grave{e}\ \{\grave{r}_i\}_{1 \leq i \leq n}$	match
		$\mathtt{cesplicede}[m;n]$	$\mathtt{splicede}\langle m, n \rangle$	spliced
CERule	$\grave{r} ::=$	$\mathtt{cerule}(p.\grave{e})$	$p \Rightarrow \grave{e}$	rule
CEPat	$\grave{p} ::=$	$\mathtt{cewildp}$	_	wildcard pattern
		$\mathtt{cefoldp}(p)$	$\mathtt{fold}(p)$	fold pattern
		$\mathtt{cetplp}[L](\{i \hookrightarrow \grave{p}_i\}_{i \in L})$	$\langle \{i \hookrightarrow \grave{p}_i\}_{i \in L} \rangle$	labeled tuple pattern
		$\mathtt{ceinp}[\ell](\grave{p})$	$\mathtt{inj}\ \ell \cdot \grave{p}$	injection pattern
		$\mathtt{cesplicedp}[m;n]$	$\mathtt{splicedp}\langle m, n \rangle$	spliced

Figure 7.3: Abstract syntax of candidate expansion types, expressions, rules and patterns in miniVerse$_\mathsf{U}^\mathbf{B}$. Candidate expansion terms are identified up to α-equivalence.

7.2.4 Syntax of Candidate Expansions

Figure 7.3 defines the syntax of candidate expansion types (or *ce-types*), $\grave{\tau}$, candidate expansion expressions (or *ce-expressions*), $\grave{e}$, candidate expansion rules (or *ce-rules*), $\grave{r}$, and candidate expansion patterns (or *ce-patterns*), $\grave{p}$. Candidate expansion terms are identified up to α-equivalence in the usual manner.

Each inner core form, except for the variable pattern form, maps onto a candidate expansion form. In particular:

- Each type form maps onto a ce-type form according to the metafunction $\mathcal{C}(\tau)$, defined in Sec. 3.2.8.

- Each expanded expression form maps onto a ce-expression form according to the metafunction $\mathcal{C}(e)$, defined as follows:

$$\mathcal{C}(x) = x$$
$$\mathcal{C}(\texttt{lam}\{\tau\}(x.e)) = \texttt{celam}\{\mathcal{C}(\tau)\}(x.\mathcal{C}(e))$$
$$\mathcal{C}(\texttt{ap}(e_1;e_2)) = \texttt{ceap}(\mathcal{C}(e_1);\mathcal{C}(e_2))$$
$$\mathcal{C}(\texttt{tlam}(t.e)) = \texttt{cetlam}(t.\mathcal{C}(e))$$
$$\mathcal{C}(\texttt{tap}\{\tau\}(e)) = \texttt{cetap}\{\mathcal{C}(\tau)\}(\mathcal{C}(e))$$
$$\mathcal{C}(\texttt{fold}\{t.\tau\}(e)) = \texttt{ceasc}\{\texttt{cerec}(t.\mathcal{C}(\tau))\}(\texttt{cefold}(\mathcal{C}(e)))$$
$$\mathcal{C}(\texttt{unfold}(e)) = \texttt{ceunfold}(\mathcal{C}(e))$$
$$\mathcal{C}(\texttt{tpl}[L](\{i \hookrightarrow e_i\}_{i \in L})) = \texttt{cetpl}[L](\{i \hookrightarrow \mathcal{C}(e_i)\}_{i \in L})$$
$$\mathcal{C}(\texttt{in}[L][\ell]\{\{i \hookrightarrow \tau_i\}_{i \in L}\}(e)) = \texttt{ceasc}\{\texttt{cesum}[L](\{i \hookrightarrow \mathcal{C}(\tau_i)\}_{i \in L})\}(\texttt{cein}[\ell](\mathcal{C}(e)))$$
$$\mathcal{C}(\texttt{match}[n]\{\tau\}(e;\{r_i\}_{1 \leq i \leq n})) = \texttt{ceasc}\{\mathcal{C}(\tau)\}(\texttt{cematch}[n](\mathcal{C}(e);\{\mathcal{C}(r_i)\}_{1 \leq i \leq n}))$$

- The expanded rule form maps onto the ce-rule form according to the metafunction $\mathcal{C}(r)$, defined as follows:

$$\mathcal{C}(\texttt{rule}(p.e)) = \texttt{cerule}(p.\mathcal{C}(e))$$

- Each expanded pattern form, except for the variable pattern form, maps onto a ce-pattern form according to the metafunction $\mathcal{C}(p)$, defined in Sec. 4.2.8.

There are three other candidate expansion forms: a ce-type form for *references to spliced unexpanded types*, $\texttt{cesplicedt}[m;n]$, a ce-expression form for *references to spliced unexpanded expressions*, $\texttt{cesplicede}[m;n]$, and a ce-pattern form for *references to spliced unexpanded patterns*, $\texttt{cesplicedp}[m;n]$.

7.2.5 Bidirectional Candidate Expansion Validation

The *bidirectional candidate expansion validation judgements* validate ce-terms and simultaneously generate their final expansions.

Judgement Form	Description
$\Delta \vdash^{\mathbb{T}} \grave{\tau} \rightsquigarrow \tau$ type	$\grave{\tau}$ is well-formed and has expansion τ assuming Δ and type splicing scene $\mathbb{T}$
$\Delta\,\Gamma \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Rightarrow \tau$	$\grave{e}$ has expansion e and synthesizes type τ assuming Δ and Γ and expression splicing scene $\mathbb{E}$
$\Delta\,\Gamma \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Leftarrow \tau$	$\grave{e}$ has expansion e when analyzed against type τ assuming Δ and Γ and expression splicing scene $\mathbb{E}$
$\Delta\,\Gamma \vdash^{\mathbb{E}} \grave{r} \rightsquigarrow r \Rightarrow \tau \mapsto \tau'$	$\grave{r}$ has expansion r and takes values of type τ to values of synthesized type τ' assuming Δ and Γ and $\mathbb{E}$
$\Delta\,\Gamma \vdash^{\mathbb{E}} \grave{r} \rightsquigarrow r \Leftarrow \tau \mapsto \tau'$	$\grave{r}$ has expansion r and takes values of type τ to values of type τ' when τ' is provided for analysis assuming Δ and Γ and $\mathbb{E}$
$\vdash^{\mathbb{P}} \grave{p} \rightsquigarrow p : \tau \dashv \mathrel{\|}^{\hat{Y}}$	$\grave{p}$ expands to p and matches values of type τ generating assumptions $\hat{Y}$ assuming pattern splicing scene $\mathbb{P}$

Expression splicing scenes, $\mathbb{E}$, are of the form $\hat{\Delta}; \hat{\Gamma}; \hat{\Psi}; \hat{\Phi}; b$, *type splicing scenes*, $\mathbb{T}$, are of the form $\hat{\Delta}; b$, and *pattern splicing scenes*, $\mathbb{P}$, are of the form $\Delta; \hat{\Phi}; b$. Their purpose is to "remember", during candidate expansion validation, the contexts, TSM environments and literal bodies from the TSM application site (cf. Rules (4.4m) and (4.6f)), because these are necessary to validate references to spliced terms. We write $\mathsf{ts}(\mathbb{E})$ for the type splicing scene constructed by dropping the unexpanded typing context and TSM environments from $\mathbb{E}$:

$$\mathsf{ts}(\hat{\Delta}; \hat{\Gamma}; \hat{\Psi}; \hat{\Phi}; b) = \hat{\Delta}; b$$

Candidate Expansion Type Validation

The *ce-type validation judgement*, $\Delta \vdash^{\mathbb{T}} \grave{\tau} \rightsquigarrow \tau$ type, is inductively defined by Rules (3.5), which were defined in Sec. 3.2.9.

Bidirectional Candidate Expansion Expression Validation

Like the bidirectionally typed expression expansion judgements, the bidirectional ce-expression validation judgements distinguish type synthesis from type analysis. The *synthetic ce-expression validation judgement*, $\Delta\,\Gamma \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Rightarrow \tau$, and the *analytic ce-expression validation judgement*, $\Delta\,\Gamma \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Leftarrow \tau$, are defined mutually inductively with Rules (7.1) and Rules (7.2) by Rules (7.6) and Rules (7.7), respectively, as follows.

Type Synthesis   Synthetic ce-expression validation is governed by the following rules.

$$\frac{}{\Delta\,\Gamma, x : \tau \vdash^{\mathbb{E}} x \rightsquigarrow x \Rightarrow \tau} \tag{7.6a}$$

$$\frac{\Delta \vdash^{\mathsf{ts}(\mathbb{E})} \grave{\tau} \rightsquigarrow \tau \text{ type} \qquad \Delta\,\Gamma \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Leftarrow \tau}{\Delta\,\Gamma \vdash^{\mathbb{E}} \mathtt{ceasc}\{\grave{\tau}\}(\grave{e}) \rightsquigarrow e \Rightarrow \tau} \tag{7.6b}$$

$$\frac{\Delta\,\Gamma \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Rightarrow \tau \qquad \Delta\,\Gamma, x:\tau \vdash^{\mathbb{E}} \grave{e}' \rightsquigarrow e' \Rightarrow \tau'}{\Delta\,\Gamma \vdash^{\mathbb{E}} \mathtt{celetval}(\grave{e}; x.\grave{e}') \rightsquigarrow \mathtt{ap}(\mathtt{lam}\{\tau\}(x.e'); e) \Rightarrow \tau'} \tag{7.6c}$$

$$\frac{\Delta \vdash^{\mathsf{ts}(\mathbb{E})} \grave{\tau}_1 \rightsquigarrow \tau_1 \text{ type} \qquad \Delta\,\Gamma, x:\tau_1 \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Rightarrow \tau_2}{\Delta\,\Gamma \vdash^{\mathbb{E}} \mathtt{celam}\{\grave{\tau}_1\}(x.\grave{e}) \rightsquigarrow \mathtt{lam}\{\tau_1\}(x.e) \Rightarrow \mathtt{parr}(\tau_1; \tau_2)} \tag{7.6d}$$

$$\frac{\Delta\,\Gamma \vdash^{\mathbb{E}} \grave{e}_1 \rightsquigarrow e_1 \Rightarrow \mathtt{parr}(\tau_2; \tau) \qquad \Delta\,\Gamma \vdash^{\mathbb{E}} \grave{e}_2 \rightsquigarrow e_2 \Leftarrow \tau_2}{\Delta\,\Gamma \vdash^{\mathbb{E}} \mathtt{ceap}(\grave{e}_1; \grave{e}_2) \rightsquigarrow \mathtt{ap}(e_1; e_2) \Rightarrow \tau} \tag{7.6e}$$

$$\frac{\Delta, t \text{ type }\Gamma \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Rightarrow \tau}{\Delta\,\Gamma \vdash^{\mathbb{E}} \mathtt{cetlam}(t.\grave{e}) \rightsquigarrow \mathtt{tlam}(t.e) \Rightarrow \mathtt{all}(t.\tau)} \tag{7.6f}$$

$$\frac{\Delta\,\Gamma \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Rightarrow \mathtt{all}(t.\tau) \qquad \Delta \vdash^{\mathsf{ts}(\mathbb{E})} \grave{\tau}' \rightsquigarrow \tau' \text{ type}}{\Delta\,\Gamma \vdash^{\mathbb{E}} \mathtt{cetap}\{\grave{\tau}'\}(\grave{e}) \rightsquigarrow \mathtt{tap}\{\tau'\}(e) \Rightarrow [\tau'/t]\tau} \tag{7.6g}$$

$$\frac{\Delta\,\Gamma \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Rightarrow \mathtt{rec}(t.\tau)}{\Delta\,\Gamma \vdash^{\mathbb{E}} \mathtt{ceunfold}(\grave{e}) \rightsquigarrow \mathtt{unfold}(e) \Rightarrow [\mathtt{rec}(t.\tau)/t]\tau} \tag{7.6h}$$

$$\frac{\{\Delta\,\Gamma \vdash^{\mathbb{E}} \grave{e}_i \rightsquigarrow e_i \Rightarrow \tau_i\}_{i\in L}}{\Delta\,\Gamma \vdash^{\mathbb{E}} \mathtt{cetpl}[L](\{i \hookrightarrow \grave{e}_i\}_{i\in L}) \rightsquigarrow \mathtt{tpl}[L](\{i \hookrightarrow e_i\}_{i\in L}) \Rightarrow \mathtt{prod}[L](\{i \hookrightarrow \tau_i\}_{i\in L})} \tag{7.6i}$$

$$\frac{\Delta\,\Gamma \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Rightarrow \mathtt{prod}[L, \ell](\{i \hookrightarrow \tau_i\}_{i\in L}; \ell \hookrightarrow \tau)}{\Delta\,\Gamma \vdash^{\mathbb{E}} \mathtt{cepr}[\ell](\grave{e}) \rightsquigarrow \mathtt{pr}[\ell](e) \Rightarrow \tau} \tag{7.6j}$$

$$\frac{n > 0 \qquad \Delta\,\Gamma \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Rightarrow \tau \qquad \{\Delta\,\Gamma \vdash^{\mathbb{E}} \grave{r}_i \rightsquigarrow r_i \Rightarrow \tau \mapsto \tau'\}_{1\le i\le n}}{\Delta\,\Gamma \vdash^{\mathbb{E}} \mathtt{cematch}[n](\grave{e}; \{\grave{r}_i\}_{1\le i\le n}) \rightsquigarrow \mathtt{match}[n]\{\tau'\}(e; \{r_i\}_{1\le i\le n}) \Rightarrow \tau'} \tag{7.6k}$$

$$\frac{\begin{array}{c}\mathsf{parseUExp}(\mathsf{subseq}(b; m; n)) = \hat{e} \qquad \langle \mathcal{D}; \Delta_{\mathsf{app}}\rangle\, \langle \mathcal{G}; \Gamma_{\mathsf{app}}\rangle \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{e} \rightsquigarrow e \Rightarrow \tau \\ \Delta \cap \Delta_{\mathsf{app}} = \varnothing \qquad \mathsf{dom}(\Gamma) \cap \mathsf{dom}(\Gamma_{\mathsf{app}}) = \varnothing\end{array}}{\Delta\,\Gamma \vdash^{\langle \mathcal{D}; \Delta_{\mathsf{app}}\rangle; \langle \mathcal{G}; \Gamma_{\mathsf{app}}\rangle; \hat{\Psi}; \hat{\Phi}; b} \mathtt{cesplicede}[m; n] \rightsquigarrow e \Rightarrow \tau} \tag{7.6l}$$

Rules (7.6a) through (7.6k) are analogous to Rules (7.1a) through (7.1k). Rule (7.6l) governs references to spliced unexpanded expressions in synthetic position, and can be understood as described in Sec. 3.2.9.

Type Analysis   Analytic ce-expression validation is governed by the following rules.

$$\frac{\Delta\,\Gamma \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Rightarrow \tau}{\Delta\,\Gamma \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Leftarrow \tau} \tag{7.7a}$$

$$\frac{\Delta\,\Gamma \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Rightarrow \tau \qquad \Delta\,\Gamma, x:\tau \vdash^{\mathbb{E}} \grave{e}' \rightsquigarrow e' \Leftarrow \tau'}{\Delta\,\Gamma \vdash^{\mathbb{E}} \mathtt{celetval}(\grave{e}; x.\grave{e}') \rightsquigarrow \mathtt{ap}(\mathtt{lam}\{\tau\}(x.e'); e) \Leftarrow \tau'} \tag{7.7b}$$

$$\frac{\Delta \; \Gamma, x : \tau_1 \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Leftarrow \tau_2}{\Delta \; \Gamma \vdash^{\mathbb{E}} \mathtt{ceanalam}(x.\grave{e}) \rightsquigarrow \mathtt{lam}\{\tau_1\}(x.e) \Leftarrow \mathtt{parr}(\tau_1; \tau_2)} \quad (7.7\mathrm{c})$$

$$\frac{\Delta, t \; \mathsf{type} \; \Gamma \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Leftarrow \tau}{\Delta \; \Gamma \vdash^{\mathbb{E}} \mathtt{cetlam}(t.\grave{e}) \rightsquigarrow \mathtt{tlam}(t.e) \Leftarrow \mathtt{all}(t.\tau)} \quad (7.7\mathrm{d})$$

$$\frac{\Delta \; \Gamma \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Leftarrow [\mathtt{rec}(t.\tau)/t]\tau}{\Delta \; \Gamma \vdash^{\mathbb{E}} \mathtt{cefold}(\grave{e}) \rightsquigarrow \mathtt{fold}\{t.\tau\}(e) \Leftarrow \mathtt{rec}(t.\tau)} \quad (7.7\mathrm{e})$$

$$\frac{\{\Delta \; \Gamma \vdash^{\mathbb{E}} \grave{e}_i \rightsquigarrow e_i \Leftarrow \tau_i\}_{i \in L}}{\Delta \; \Gamma \vdash^{\mathbb{E}} \mathtt{cetpl}[L](\{i \hookrightarrow \grave{e}_i\}_{i \in L}) \rightsquigarrow \mathtt{tpl}[L](\{i \hookrightarrow e_i\}_{i \in L}) \Leftarrow \mathtt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L})}$$
$$(7.7\mathrm{f})$$

$$\left(\frac{\Delta \; \Gamma \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Leftarrow \tau}{\begin{array}{c} \Delta \; \Gamma \vdash^{\mathbb{E}} \underset{\rightsquigarrow}{\mathtt{cein}[\ell](\grave{e})} \\ \mathtt{in}[L, \ell][\ell]\{\{i \hookrightarrow \tau_i\}_{i \in L}; \ell \hookrightarrow \tau\}(e) \Leftarrow \mathtt{sum}[L, \ell](\{i \hookrightarrow \tau_i\}_{i \in L}; \ell \hookrightarrow \tau) \end{array}} \right) \quad (7.7\mathrm{g})$$

$$\frac{\Delta \; \Gamma \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Rightarrow \tau \qquad \{\Delta \; \Gamma \vdash^{\mathbb{E}} \grave{r}_i \rightsquigarrow r_i \Leftarrow \tau \mapsto \tau'\}_{1 \le i \le n}}{\Delta \; \Gamma \vdash^{\mathbb{E}} \mathtt{cematch}[n](\grave{e}; \{\grave{r}_i\}_{1 \le i \le n}) \rightsquigarrow \mathtt{match}[n]\{\tau'\}(e; \{r_i\}_{1 \le i \le n}) \Leftarrow \tau'} \quad (7.7\mathrm{h})$$

$$\frac{\begin{array}{c} \mathsf{parseUExp}(\mathsf{subseq}(b; m; n)) = \hat{e} \qquad \langle \mathcal{D}; \Delta_{\mathrm{app}} \rangle \; \langle \mathcal{G}; \Gamma_{\mathrm{app}} \rangle \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{e} \rightsquigarrow e \Leftarrow \tau \\ \Delta \cap \Delta_{\mathrm{app}} = \varnothing \qquad \mathrm{dom}(\Gamma) \cap \mathrm{dom}(\Gamma_{\mathrm{app}}) = \varnothing \end{array}}{\Delta \; \Gamma \vdash^{\langle \mathcal{D}; \Delta_{\mathrm{app}} \rangle; \langle \mathcal{G}; \Gamma_{\mathrm{app}} \rangle; \hat{\Psi}; \hat{\Phi}; b} \mathtt{cesplicede}[m; n] \rightsquigarrow e \Leftarrow \tau} \quad (7.7\mathrm{i})$$

Rules (7.7a) through (7.7h) are analagous to Rules (7.2a) through (7.2h). Rule (7.7i) governs references to spliced unexpanded expressions in analytic position.

Bidirectional Candidate Expansion Rule Validation

The *synthetic ce-rule validation judgement* is defined mutually inductively with Rules (7.1) by the following rule.

$$\frac{\Delta \vdash p : \tau \dashv Y \qquad \Delta \; \Gamma \cup Y \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Rightarrow \tau'}{\Delta \; \Gamma \vdash^{\mathbb{E}} \mathtt{cerule}(p.\grave{e}) \rightsquigarrow \mathtt{rule}(p.e) \Rightarrow \tau \mapsto \tau'} \quad (7.8)$$

The *analytic ce-rule validation judgement* is defined mutually inductively with Rules (7.2) by the following rule.

$$\frac{\Delta \vdash p : \tau \dashv Y \qquad \Delta \; \Gamma \cup Y \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Leftarrow \tau'}{\Delta \; \Gamma \vdash^{\mathbb{E}} \mathtt{cerule}(p.\grave{e}) \rightsquigarrow \mathtt{rule}(p.e) \Leftarrow \tau \mapsto \tau'} \quad (7.9)$$

Candidate Expansion Pattern Validation

The *ce-pattern validation judgement* is inductively defined by the following rules, which are written identically to Rules (4.9).

$$\frac{}{\vdash^{\mathbb{P}} \texttt{cewildp} \leadsto \texttt{wildp} : \tau \dashv\!\!\mid^{\langle\varnothing;\varnothing\rangle}} \tag{7.10a}$$

$$\frac{\vdash^{\mathbb{P}} \grave{p} \leadsto p : [\texttt{rec}(t.\tau)/t]\tau \dashv\!\!\mid^{\hat{Y}}}{\vdash^{\mathbb{P}} \texttt{cefoldp}(\grave{p}) \leadsto \texttt{foldp}(p) : \texttt{rec}(t.\tau) \dashv\!\!\mid^{\hat{Y}}} \tag{7.10b}$$

$$\frac{\{\vdash^{\mathbb{P}} \grave{p}_i \leadsto p_i : \tau_i \dashv\!\!\mid^{\hat{Y}_i}\}_{i \in L}}{\left(\begin{array}{c} \vdash^{\mathbb{P}} \texttt{cetplp}[L](\{i \hookrightarrow \grave{p}_i\}_{i \in L}) \\ \leadsto \\ \texttt{tplp}[L](\{i \hookrightarrow p_i\}_{i \in L}) : \texttt{prod}[L](\{i \hookrightarrow \tau_i\}_{i \in L}) \dashv\!\!\mid^{\cup_{i \in L} \hat{Y}_i} \end{array}\right)} \tag{7.10c}$$

$$\frac{\vdash^{\mathbb{P}} \grave{p} \leadsto p : \tau \dashv\!\!\mid^{\hat{Y}}}{\vdash^{\mathbb{P}} \texttt{ceinp}[\ell](\grave{p}) \leadsto \texttt{inp}[\ell](p) : \texttt{sum}[L,\ell](\{i \hookrightarrow \tau_i\}_{i \in L}; \ell \hookrightarrow \tau) \dashv\!\!\mid^{\hat{Y}}} \tag{7.10d}$$

$$\frac{\texttt{parseUPat}(\texttt{subseq}(b;m;n)) = \hat{p} \qquad \Delta \vdash_{\hat{\Phi}} \hat{p} \leadsto p : \tau \dashv\!\!\mid \hat{Y}}{\vdash^{\Delta;\hat{\Phi};b} \texttt{cesplicedp}[m;n] \leadsto p : \tau \dashv\!\!\mid^{\hat{Y}}} \tag{7.10e}$$

Candidate Expansion Expressibility

The following lemma establishes that each well-typed expanded expression, *e*, can be expressed as a valid ce-expression, $\mathcal{C}(e)$, that synthesizes the same type under the same contexts and any expression splicing scene.

Theorem 7.4 (Candidate Expansion Expression Expressibility)**.** *Both of the following hold:*

1. *If $\Delta \, \Gamma \vdash e : \tau$ then $\Delta \, \Gamma \vdash^{\mathbb{E}} \mathcal{C}(e) \leadsto e \Rightarrow \tau$.*
2. *If $\Delta \, \Gamma \vdash r : \tau \mapsto \tau'$ then $\Delta \, \Gamma \vdash^{\mathbb{E}} \mathcal{C}(r) \leadsto r \Rightarrow \tau \mapsto \tau'$.*

Proof. By mutual rule induction over Rules (4.1) and Rule (4.2). In each case, we apply the IH, part 1 to or over each typing premise, the IH, part 2 over each rule typing premise, Lemma 3.18 to or over each type formation premise and then derive the conclusion by applying Rules (7.6) and Rule (7.8) as needed. $\square$

The following lemma establishes that every well-typed expanded pattern that generates no hypotheses can be expressed as a ce-pattern.

Lemma 7.5 (Candidate Expansion Pattern Expressibility)**.** *If $\Delta \vdash p : \tau \dashv\!\!\mid \varnothing$ then $\vdash^{\Delta;\hat{\Phi};b} \mathcal{C}(p) \leadsto p : \tau \dashv\!\!\mid^{\langle\varnothing;\varnothing\rangle}$.*

Proof. The proof is nearly identical to the proof of Lemma 4.20, differing only in that each mention of a rule in Rules (4.9) is replaced by a mention of the corresponding rule in Rules (7.10). $\square$

7.2.6 Metatheory

The following theorem establishes that typed pattern expansion produces an expanded pattern that matches values of the specified type and generates the same hypotheses. It must be stated mutually with the corresponding theorem about candidate expansion patterns, because the judgements are mutually defined.

Theorem 7.6 (Typed Pattern Expansion). *Both of the following hold:*

1. *If* $\Delta \vdash_{\langle \mathcal{A};\Phi;\mathcal{I}\rangle} \hat{p} \rightsquigarrow p : \tau \dashv \langle \mathcal{G};Y\rangle$ *then* $\Delta \vdash p : \tau \dashv Y$.
2. *If* $\vdash^{\Delta;\langle \mathcal{A};\Phi;\mathcal{I}\rangle;b} \hat{p} \rightsquigarrow p : \tau \dashv^{\langle \mathcal{G};Y\rangle}$ *then* $\Delta \vdash p : \tau \dashv Y$.

Proof. My mutual rule induction over Rules (7.5) and Rules (7.10).

1. We induct on the premise. In the following, let $\hat{Y} = \langle \mathcal{G};Y\rangle$ and $\hat{\Phi} = \langle \mathcal{A};\Phi;\mathcal{I}\rangle$.
   Case (7.5a) through (7.5f). In each of these cases, the proof is written identically to the proof of the corresponding case in the proof of Theorem 4.21.
   Case (7.5g). We have:

(1)	$\hat{p} = \mathtt{uplit}[b]$	by assumption
(2)	$\Phi = \Phi', a \hookrightarrow \mathtt{uptsm}(\tau; e_{\mathrm{parse}})$	by assumption
(3)	$\mathcal{I} = \mathcal{I}', \tau \hookrightarrow a$	by assumption
(4)	$b \downarrow e_{\mathrm{body}}$	by assumption
(5)	$e_{\mathrm{parse}}(e_{\mathrm{body}}) \Downarrow \mathtt{inj\ Success} \cdot e_{\mathrm{cand}}$	by assumption
(6)	$e_{\mathrm{cand}} \uparrow_{\mathsf{CEPat}} \hat{p}$	by assumption
(7)	$\vdash^{\Delta;\langle \mathcal{A};\Phi',a\hookrightarrow\mathtt{uptsm}(\tau;e_{\mathrm{parse}});\mathcal{I}',\tau\hookrightarrow a\rangle;b} \hat{p} \rightsquigarrow p : \tau \dashv^{\langle \mathcal{G};Y\rangle}$	by assumption
(8)	$\Delta \vdash p : \tau \dashv Y$	by IH, part 2 on (7)

2. We induct on the premise. In the following, let $\hat{Y} = \langle \mathcal{G};Y\rangle$ and $\hat{\Phi} = \langle \mathcal{A};\Phi;\mathcal{I}\rangle$.
   Case (7.10a) through (7.10e). In each case, the proof is written identically to the proof of the corresponding case in the proof of Theorem 4.21.

The mutual induction can be shown to be well-founded by showing that the following numeric metric on the judgements that we induct on is decreasing:

$$\|\Delta \vdash_{\hat{\Phi}} \hat{p} \rightsquigarrow p : \tau \dashv \hat{Y}\| = \|\hat{p}\|$$
$$\|\vdash^{\Delta;\hat{\Phi};b} \hat{p} \rightsquigarrow p : \tau \dashv^{\hat{Y}}\| = \|b\|$$

where $\|b\|$ is the length of b and $\|\hat{p}\|$ is the sum of the lengths of the literal bodies in $\hat{p}$,

$$\|\hat{x}\| = 0$$
$$\|\mathtt{ufoldp}(\hat{p})\| = \|\hat{p}\|$$
$$\|\mathtt{utplp}[L](\{i \hookrightarrow \hat{p}_i\}_{i\in L})\| = \sum_{i\in L}\|\hat{p}_i\|$$
$$\|\mathtt{uinp}[\ell](\hat{p})\| = \|\hat{p}\|$$
$$\|\mathtt{uapuptsm}[b][\hat{a}]\| = \|b\|$$
$$\|\mathtt{uplit}[b]\| = \|b\|$$

The only case in the proof of part 1 that invokes part 2 are Case (7.5f) and (7.5g). There, we have that the metric remains stable:

$$\|\Delta \vdash_{\hat{\Phi}, \hat{a} \leadsto a \hookrightarrow \mathtt{uetsm}(\tau; e_{\text{parse}})} \mathtt{uapuptsm}[b][\hat{a}] \leadsto p : \tau \dashv \hat{Y}\|$$

$$= \|\Delta \vdash_{\langle \mathcal{A}; \Phi', a \hookrightarrow \mathtt{uptsm}(\tau; e_{\text{parse}}); \mathcal{I}', \tau \hookrightarrow a \rangle} \mathtt{uplit}[b] \leadsto p : \tau \dashv \hat{Y}\|$$

$$= \|\vdash^{\Delta; \hat{\Phi}, \hat{a} \leadsto a \hookrightarrow \mathtt{uetsm}(\tau; e_{\text{parse}}); b} \hat{p} \leadsto p : \tau \dashv^{\hat{Y}}\|$$

$$= \|b\|$$

The only case in the proof of part 2 that invokes part 1 is Case (7.10e). There, we have that $\mathsf{parseUPat}(\mathtt{subseq}(b; m; n)) = \hat{p}$ and the IH is applied to the judgement $\Delta \vdash_{\hat{\Phi}} \hat{p} \leadsto p : \tau \dashv \hat{Y}$. Because the metric is stable when passing from part 1 to part 2, we must have that it is strictly decreasing in the other direction:

$$\|\Delta \vdash_{\hat{\Phi}} \hat{p} \leadsto p : \tau \dashv \hat{Y}\| < \|\vdash^{\Delta; \hat{\Phi}; b} \mathtt{cesplicedp}[m; n] \leadsto p : \tau \dashv^{\hat{Y}}\|$$

i.e. by the definitions above,

$$\|\hat{p}\| < \|b\|$$

This is established by appeal to Condition 3.22, which states that subsequences of b are no longer than b, and the following condition, which states that an unexpanded pattern constructed by parsing a textual sequence b is strictly smaller, as measured by the metric defined above, than the length of b, because some characters must necessarily be used to delimit each literal body.

Condition 7.7 (Pattern Parsing Monotonicity). *If* $\mathsf{parseUPat}(b) = \hat{p}$ *then* $\|\hat{p}\| < \|b\|$.

Combining Conditions 3.22 and 7.7, we have that $\|\hat{e}\| < \|b\|$ as needed. $\square$

Finally, the following theorem establishes that bidirectionally typed expression and rule expansion produces expanded expressions and rules of the appropriate type under the appropriate contexts. These statements must be stated mutually with the corresponding statements about birectional ce-expression and ce-rule validation because the judgements are mutually defined.

Theorem 7.8 (Typed Expansion). *Letting* $\hat{\Psi} = \langle \mathcal{A}; \Psi; \mathcal{I} \rangle$, *if* $\Delta \vdash \Psi$ ueTSMs *then all of the following hold:*

1. (a)
   i. *If* $\langle \mathcal{D}; \Delta \rangle \langle \mathcal{G}; \Gamma \rangle \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{e} \leadsto e \Rightarrow \tau$ *then* $\Delta \, \Gamma \vdash e : \tau$.
   ii. *If* $\langle \mathcal{D}; \Delta \rangle \langle \mathcal{G}; \Gamma \rangle \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{r} \leadsto r \Rightarrow \tau \mapsto \tau'$ *then* $\Delta \, \Gamma \vdash r : \tau \mapsto \tau'$.
   (b)
   i. *If* $\langle \mathcal{D}; \Delta \rangle \langle \mathcal{G}; \Gamma \rangle \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{e} \leadsto e \Leftarrow \tau$ *and* $\Delta \vdash \tau$ type *then* $\Delta \, \Gamma \vdash e : \tau$.
   ii. *If* $\langle \mathcal{D}; \Delta \rangle \langle \mathcal{G}; \Gamma \rangle \vdash_{\hat{\Psi}; \hat{\Phi}} \hat{r} \leadsto r \Leftarrow \tau \mapsto \tau'$ *and* $\Delta \vdash \tau'$ type *then* $\Delta \, \Gamma \vdash r : \tau \mapsto \tau'$.
2. (a)
   i. *If* $\Delta \, \Gamma \vdash^{\langle \mathcal{D}; \Delta_{app} \rangle; \langle \mathcal{G}; \Gamma_{app} \rangle; \hat{\Psi}; \hat{\Phi}; b} \hat{e} \leadsto e \Rightarrow \tau$ *and* $\Delta \cap \Delta_{app} = \emptyset$ *and* $dom(\Gamma) \cap dom(\Gamma_{app}) = \emptyset$ *then* $\Delta \cup \Delta_{app} \, \Gamma \cup \Gamma_{app} \vdash e : \tau$.
   ii. *If* $\Delta \, \Gamma \vdash^{\langle \mathcal{D}; \Delta_{app} \rangle; \langle \mathcal{G}; \Gamma_{app} \rangle; \hat{\Psi}; \hat{\Phi}; b} \hat{r} \leadsto r \Rightarrow \tau \mapsto \tau'$ *and* $\Delta \cap \Delta_{app} = \emptyset$ *and* $dom(\Gamma) \cap dom(\Gamma_{app}) = \emptyset$ *then* $\Delta \cup \Delta_{app} \, \Gamma \cup \Gamma_{app} \vdash r : \tau \mapsto \tau'$.

(b)   i. *If* $\Delta \Gamma \vdash^{\langle \mathcal{D};\Delta_{app}\rangle;\langle \mathcal{G};\Gamma_{app}\rangle;\hat{\Psi};\hat{\Phi};b} \grave{e} \rightsquigarrow e \Leftarrow \tau$ *and* $\Delta \cap \Delta_{app} = \varnothing$ *and* $dom(\Gamma) \cap$
    $dom(\Gamma_{app}) = \varnothing$ *and* $\Delta \cup \Delta_{app} \vdash \tau$ type *then* $\Delta \cup \Delta_{app} \Gamma \cup \Gamma_{app} \vdash e : \tau$.
   ii. *If* $\Delta \Gamma \vdash^{\langle \mathcal{D};\Delta_{app}\rangle;\langle \mathcal{G};\Gamma_{app}\rangle;\hat{\Psi};\hat{\Phi};b} \grave{r} \rightsquigarrow r \Leftarrow \tau \Rightarrow \tau'$ *and* $\Delta \cap \Delta_{app} = \varnothing$ *and*
    $dom(\Gamma) \cap dom(\Gamma_{app}) = \varnothing$ *and* $\Delta \cup \Delta_{app} \vdash \tau'$ type *then* $\Delta \cup \Delta_{app} \Gamma \cup \Gamma_{app} \vdash r :$
    $\tau \Rightarrow \tau'$.

Proof. By mutual rule induction over Rules (7.1), Rules (7.2), Rule (7.3), Rule (7.4), Rules (7.6), Rules (7.7), Rule (7.8) and Rule (7.9). In the following, we refer to the induction hypothesis applied to the assumption $\Delta \vdash \Psi$ ueTSMs as simply the "IH". When we apply the induction hypothesis to a different argument, we refer to it as the "Outer IH".

1. In the following, let $\hat{\Delta} = \langle \mathcal{D};\Delta \rangle$ and $\hat{\Gamma} = \langle \mathcal{G};\Gamma \rangle$. We have:
   (a)   i. We induct on the assumption.
      Case (7.1a). We have:

(1)	$e = x$	by assumption
(2)	$\Gamma = \Gamma', x : \tau$	by assumption
(3)	$\Delta \Gamma', x : \tau \vdash x : \tau$	by Rule (4.1a)

      Case (7.1b). We have:

(1)	$\hat{e} = \mathtt{uasc}\{\hat{\tau}\}(\hat{e}')$	by assumption
(2)	$\hat{\Delta} \vdash \hat{\tau} \rightsquigarrow \tau$ type	by assumption
(3)	$\hat{\Delta} \hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e}' \rightsquigarrow e \Leftarrow \tau$	by assumption
(4)	$\Delta \vdash \tau$ type	by Lemma 3.12 on (2)
(5)	$\Delta \Gamma \vdash e : \tau$	by IH, part 1(b)(i) to (3) and (4)

      Case (7.1c) through (7.1k). In each of these cases, we apply:
      - Lemma 3.12 to or over all type expansion premises.
      - The IH, part 1(a)(i) to or over all synthetic typed expression expansion premises.
      - The IH, part 1(a)(ii) to or over all synthetic rule expansion premises.
      - The IH, part 1(b)(i) to or over all analytic typed expression expansion premises.

      We then derive the conclusion by applying Rules (4.1) and Rule (4.2) as needed.

      Case (7.1l). We have:

(1)	$\hat{e} = \mathtt{usyntaxue}\{e_{\mathrm{parse}}\}\{\hat{\tau}'\}(\hat{a}.\hat{e}')$	by assumption
(2)	$\hat{\Delta} \vdash \hat{\tau}' \rightsquigarrow \tau'$ type	by assumption
(3)	$\varnothing \varnothing \vdash e_{\mathrm{parse}} : \mathtt{parr}(\mathtt{Body};\mathtt{ParseResultExp})$	
		by assumption
(4)	$\hat{\Delta} \hat{\Gamma} \vdash_{\langle \mathcal{A}\uplus\hat{a}\rightsquigarrow a;\Psi,a\hookrightarrow\mathtt{uetsm}(\tau';e_{\mathrm{parse}});\mathcal{I}\rangle;\hat{\Phi}} \hat{e}' \rightsquigarrow e \Rightarrow \tau$	
		by assumption
(5)	$\Delta \vdash \Psi$ ueTSMs	by assumption
(6)	$\Delta \vdash \tau'$ type	by Lemma 3.12 to (2)

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(7) $\Delta \vdash \Psi, \hat{a} \hookrightarrow \texttt{uetsm}(\tau'; e_{\mathrm{parse}})$ ueTSMs — by Definition 4.14 on (5), (6) and (3)

(8) $\Delta \, \Gamma \vdash e : \tau$ — by Outer IH, part 1(a)(i) on (7) and (4)

Case (7.1m). We have:

(1) $\hat{e} = \texttt{uapuetsm}[b][\hat{a}]$ — by assumption

(2) $\hat{\Psi} = \langle \mathcal{A}' \uplus \hat{a} \rightsquigarrow a; \Psi', a \hookrightarrow \texttt{uetsm}(\tau; e_{\mathrm{parse}}); \mathcal{I} \rangle$ — by assumption

(3) $b \downarrow e_{\mathrm{body}}$ — by assumption

(4) $e_{\mathrm{parse}}(e_{\mathrm{body}}) \Downarrow \texttt{inj Success} \cdot e_{\mathrm{cand}}$ — by assumption

(5) $e_{\mathrm{cand}} \uparrow_{\mathsf{CEExp}} \grave{e}$ — by assumption

(6) $\varnothing \, \varnothing \vdash^{\hat{\Delta}; \hat{\Gamma}; \hat{\Psi}; \hat{\Phi}; b} \grave{e} \rightsquigarrow e \Leftarrow \tau$ — by assumption

(7) $\Delta \vdash \Psi$ ueTSMs — by assumption

(8) $\Delta \vdash \tau$ type — by Definition 4.14 on (7)

(9) $\varnothing \cap \Delta = \varnothing$ — by finite set intersection identity

(10) $\varnothing \cap \mathrm{dom}(\Gamma) = \varnothing$ — by finite set intersection identity

(11) $\varnothing \cup \Delta \, \varnothing \cup \Gamma \vdash e : \tau$ — by IH, part 2(a)(i) on (6), (9), (10) and (8)

(12) $\Delta \, \Gamma \vdash e : \tau$ — by definition of finite set and finite function union over (11)

Case (7.1n). We have:

(1) $\hat{e} = \texttt{uimplicite}[\hat{a}](\hat{e})$ — by assumption

(2) $\hat{\Psi} = \langle \mathcal{A}' \uplus \hat{a} \rightsquigarrow a; \Psi', a \hookrightarrow \texttt{uetsm}(\tau'; e_{\mathrm{parse}}); \mathcal{I} \rangle$ — by assumption

(3) $\hat{\Delta} \, \hat{\Gamma} \vdash_{\langle \mathcal{A}' \uplus \hat{a} \rightsquigarrow a; \Psi', a \hookrightarrow \texttt{uetsm}(\tau'; e_{\mathrm{parse}}); \mathcal{I} \uplus \tau \hookrightarrow a \rangle; \hat{\Phi}} \hat{e} \rightsquigarrow e \Rightarrow \tau$ — by assumption

(4) $\Delta \, \Gamma \vdash e : \tau$ — by IH, part 1(a)(i) on (3)

Case (7.1o). We have:

(1) $\hat{e} = \texttt{usyntaxup}\{e_{\mathrm{parse}}\}\{\hat{\tau}'\}(\hat{a}.\hat{e}')$ — by assumption

(2) $\hat{\Delta} \vdash \hat{\tau}' \rightsquigarrow \tau'$ type — by assumption

(3) $\hat{\Delta} \, \hat{\Gamma} \vdash_{\Psi; \hat{\Phi}, \hat{a} \rightsquigarrow a \hookrightarrow \texttt{uptsm}(\tau'; e_{\mathrm{parse}})} \hat{e}' \rightsquigarrow e \Rightarrow \tau$ — by assumption

(4) $\Delta \, \Gamma \vdash e : \tau$ — by IH, part 1(a)(i) on (3)

Case (7.1p). We have:

(1) $\hat{e} = \texttt{uimplicitp}[\hat{a}](\hat{e})$ — by assumption

(2) $\hat{\Phi} = \langle \mathcal{A} \uplus \hat{a} \rightsquigarrow a; \Phi, a \hookrightarrow \texttt{uptsm}(\tau'; e_{\mathrm{parse}}); \mathcal{I} \rangle$

$$\text{(3)}\ \hat{\Delta}\ \hat{\Gamma} \vdash_{\hat{\Psi};\langle \mathcal{A}\uplus\hat{a}\leadsto a;\Phi,a\hookrightarrow\texttt{uptsm}(\tau';e_{\text{parse}});\mathcal{I}\uplus\tau\hookrightarrow a\rangle}\ \hat{e} \leadsto e \Rightarrow \tau \qquad \text{by assumption}$$

$$\text{(4)}\ \Delta\ \Gamma \vdash e : \tau \qquad \text{by IH, part 1(a)(i) on (3)}$$

ii. We induct on the assumption. There is one case.

Case (7.3). We have:

(1)	$\hat{r} = \texttt{urule}(\hat{p}.\hat{e})$	by assumption
(2)	$r = \texttt{rule}(p.e)$	by assumption
(3)	$\Delta \vdash_{\hat{\Phi}} \hat{p} \leadsto p : \tau \dashv\!\vdash \langle \mathcal{A}';Y\rangle$	by assumption
(4)	$\hat{\Delta}\ \langle\mathcal{A}\uplus\mathcal{A}';\Gamma\cup Y\rangle \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \leadsto e \Rightarrow \tau'$	by assumption
(5)	$\Delta \vdash p : \tau \dashv\!\vdash Y$	by Theorem 7.6, part 1 on (3)
(6)	$\Delta\ \Gamma\cup Y \vdash e : \tau'$	by IH, part 1(a)(i) on (4)
(7)	$\Delta\ \Gamma \vdash \texttt{rule}(p.e) : \tau \Mapsto \tau'$	by Rule (4.2) on (5) and (6)

(b)   i. We induct on the assumption.

Case (7.2a). We have:

(1)	$\hat{\Delta}\ \hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \leadsto e \Rightarrow \tau$	by assumption
(2)	$\Delta\ \Gamma \vdash e : \tau$	by IH, part 1(a)(i) on (1)

Case (7.2b) through **(7.2h).** In each of these cases, we apply:

- Lemma 3.12 to or over all type expansion premises.
- The IH, part 1(a)(i) to or over all synthetic typed expression expansion premises.
- The IH, part 1(a)(ii) to or over all synthetic rule expansion premises.
- The IH, part 1(b)(i) to or over all analytic typed expression expansion premises.

We then derive the conclusion by applying Rules (4.1) and Rule (4.2) as needed.

Case (7.2i). We have:

(3)	$\hat{e} = \texttt{usyntaxue}\{e_{\text{parse}}\}\{\hat{\tau}'\}(\hat{a}.\hat{e}')$	by assumption
(4)	$\hat{\Delta} \vdash \hat{\tau}' \leadsto \tau'\ \text{type}$	by assumption
(5)	,	by assumption
(6)	$\hat{\Delta}\ \hat{\Gamma} \vdash_{\hat{\Psi},\hat{a}\leadsto a\hookrightarrow\texttt{uetsm}(\tau';e_{\text{parse}});\hat{\Phi}}\ \hat{e}' \leadsto e \Leftarrow \tau$	by assumption
(7)	$\Delta \vdash \Psi\ \text{ueTSMs}$	by assumption
(8)	$\Delta \vdash \tau'\ \text{type}$	by Lemma 3.12 to (4)
(9)	$\Delta \vdash \Psi,\hat{a} \hookrightarrow \texttt{uetsm}(\tau';e_{\text{parse}})\ \text{ueTSMs}$	by Definition 4.14 on (7), (8) and (5)

(10) $\Delta\,\Gamma \vdash e : \tau$ by IH, part 1(b)(i) on (6)

Case (7.2j). We have:

(1) $\hat{e} = \texttt{uapuetsm}[b][\hat{a}]$ by assumption

(2) $\hat{\Psi} = \hat{\Psi}', \hat{a} \rightsquigarrow a \hookrightarrow \texttt{uetsm}(\tau; e_{\text{parse}})$ by assumption

(3) $b \downarrow e_{\text{body}}$ by assumption

(4) $e_{\text{parse}}(e_{\text{body}}) \Downarrow \texttt{inj Success} \cdot e_{\text{cand}}$ by assumption

(5) $e_{\text{cand}} \uparrow_{\text{CEExp}} \grave{e}$ by assumption

(6) $\varnothing\,\varnothing \vdash^{\hat{\Delta};\hat{\Gamma};\hat{\Psi};\hat{\Phi};b} \grave{e} \rightsquigarrow e \Leftarrow \tau$ by assumption

(7) $\varnothing \cap \Delta = \varnothing$ by finite set intersection identity

(8) $\varnothing \cap \text{dom}(\Gamma) = \varnothing$ by finite set intersection identity

(9) $\varnothing \cup \Delta\,\varnothing \cup \Gamma \vdash e : \tau$ by IH, part 2(b)(i) on (6), (7), and (8)

(10) $\Delta\,\Gamma \vdash e : \tau$ by definition of finite set union over (9)

Case (7.2k). We have:

(1) $\hat{e} = \texttt{uelit}[b]$ by assumption

(2) $\hat{\Psi} = \langle \mathcal{A}; \Psi, a \hookrightarrow \texttt{uetsm}(\tau; e_{\text{parse}}); \mathcal{I} \uplus \tau \hookrightarrow a \rangle$ by assumption

(3) $b \downarrow e_{\text{body}}$ by assumption

(4) $e_{\text{parse}}(e_{\text{body}}) \Downarrow \texttt{inj Success} \cdot e_{\text{cand}}$ by assumption

(5) $e_{\text{cand}} \uparrow_{\text{CEExp}} \grave{e}$ by assumption

(6) $\varnothing\,\varnothing \vdash^{\hat{\Delta};\hat{\Gamma};\langle \mathcal{A};\Psi,a\hookrightarrow\texttt{uetsm}(\tau;e_{\text{parse}});\mathcal{I}\uplus\tau\hookrightarrow a\rangle;\hat{\Phi};b} \grave{e} \rightsquigarrow e \Leftarrow \tau$ by assumption

(7) $\varnothing \cap \Delta = \varnothing$ by finite set intersection identity

(8) $\varnothing \cap \text{dom}(\Gamma) = \varnothing$ by finite set intersection identity

(9) $\varnothing \cup \Delta\,\varnothing \cup \Gamma \vdash e : \tau$ by IH, part 2(a)(i) on (6), (7), and (8)

(10) $\Delta\,\Gamma \vdash e : \tau$ by definition of finite set union over (9)

Case (7.2l). We have:

(1) $\hat{e} = \texttt{usyntaxup}\{e_{\text{parse}}\}\{\hat{\tau}'\}(\hat{a}.\hat{e}')$ by assumption

(2) $\hat{\Delta} \vdash \hat{\tau}' \rightsquigarrow \tau' \text{ type}$ by assumption

(3) $\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi},\hat{a}\rightsquigarrow a\hookrightarrow\texttt{uptsm}(\tau';e_{\text{parse}})} \hat{e}' \rightsquigarrow e \Leftarrow \tau$ by assumption

(4) $\Delta\,\Gamma \vdash e : \tau$ by IH, part 1(b)(i) on (3)

Case (7.2m). We have:

    (1) $\hat{e} = \texttt{uimplicitp}[\hat{a}](\hat{e})$ by assumption

    (2) $\hat{\Phi} = \langle \mathcal{A} \uplus \hat{a} \rightsquigarrow a; \Phi, a \hookrightarrow \texttt{uptsm}(\tau'; e_{\text{parse}}); \mathcal{I} \rangle$

                                               by assumption

    (3) $\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\langle \mathcal{A}\uplus\hat{a}\rightsquigarrow a;\Phi,a\hookrightarrow\texttt{uptsm}(\tau';e_{\text{parse}});\mathcal{I}\uplus\tau\hookrightarrow a\rangle} \hat{e} \rightsquigarrow e \Leftarrow \tau$

                                               by assumption

    (4) $\Delta\,\Gamma \vdash e : \tau$ by IH, part 1(b)(i) on (3)

ii. We induct on the assumption. There is one case.

Case (7.4). We have:

    (1) $\hat{r} = \texttt{urule}(\hat{p}.\hat{e})$ by assumption

    (2) $r = \texttt{rule}(p.e)$ by assumption

    (3) $\Delta \vdash_{\hat{\Phi}} \hat{p} \rightsquigarrow p : \tau \dashv\!\!| \langle \mathcal{A}'; Y \rangle$ by assumption

    (4) $\hat{\Delta}\,\langle \mathcal{A} \uplus \mathcal{A}'; \Gamma \cup Y \rangle \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \rightsquigarrow e \Leftarrow \tau'$ by assumption

    (5) $\Delta \vdash p : \tau \dashv\!\!| Y$ by Theorem 7.6, part 1 on (3)

    (6) $\Delta\,\Gamma \cup Y \vdash e : \tau'$ by IH, part 1(b)(i) on (4)

    (7) $\Delta\,\Gamma \vdash \texttt{rule}(p.e) : \tau \mapsto \tau'$ by Rule (4.2) on (5) and (6)

2. In the following, let $\hat{\Delta} = \langle \mathcal{D}; \Delta_{\text{app}} \rangle$ and $\hat{\Gamma} = \langle \mathcal{G}; \Gamma_{\text{app}} \rangle$ and $\mathbb{E} = \hat{\Delta}; \hat{\Gamma}; \hat{\Psi}; \hat{\Phi}; b$.

  (a) i. We induct on the assumption.

Case (7.6a). We have:

    (1) $e = x$ by assumption

    (2) $\Gamma = \Gamma', x : \tau$ by assumption

    (3) $\Delta\,\Gamma', x : \tau \vdash x : \tau$ by Rule (4.1a)

Case (7.6b). We have:

    (1) $\grave{e} = \texttt{ceasc}\{\grave{\tau}\}(\grave{e}')$ by assumption

    (2) $\Delta \cap \Delta_{\text{app}} = \varnothing$ by assumption

    (3) $\text{dom}(\Gamma) \cap \text{dom}(\Gamma_{\text{app}}) = \varnothing$ by assumption

    (4) $\Delta \vdash^{\text{ts}(\mathbb{E})} \grave{\tau} \rightsquigarrow \tau\ \text{type}$ by assumption

    (5) $\Delta\,\Gamma \vdash^{\mathbb{E}} \grave{e}' \rightsquigarrow e \Leftarrow \tau$ by assumption

    (6) $\Delta \cup \Delta_{\text{app}} \vdash \tau\ \text{type}$ by Lemma 3.19 on (4)

    (7) $\Delta\,\Gamma \vdash e : \tau$ by IH, part 2(b)(i) to (5), (2), (3) and (6)

Case (7.6c) through (7.6k). In each of these cases, we apply:

- Lemma 3.19 to or over all ce-type validation premises.
- The IH, part 2(a)(i) to or over all synthetic ce-expression validation premises.
- The IH, part 2(a)(ii) to or over all synthetic ce-rule validation premises.

- The IH, part 2(b)(i) to or over all analytic ce-expression validation premises.

We then derive the conclusion by applying Rules (4.1), Rule (4.2), Lemma 4.1, the identification convention and exchange as needed.

Case (7.6l). We have:

(1)	$\grave{e} = \mathtt{cesplicede}[m; n]$	by assumption
(2)	$\mathsf{parseUExp}(\mathsf{subseq}(b; m; n)) = \hat{e}$	by assumption
(3)	$\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \rightsquigarrow e \Rightarrow \tau$	by assumption
(4)	$\Delta \cap \Delta_{\mathsf{app}} = \varnothing$	by assumption
(5)	$\mathsf{dom}(\Gamma) \cap \mathsf{dom}(\Gamma_{\mathsf{app}}) = \varnothing$	by assumption
(6)	$\Delta_{\mathsf{app}}\,\Gamma_{\mathsf{app}} \vdash e : \tau$	by IH, part 1(a)(i) on (3)
(7)	$\Delta \cup \Delta_{\mathsf{app}}\,\Gamma \cup \Gamma_{\mathsf{app}} \vdash e : \tau$	by Lemma 4.1 over Δ and Γ and exchange on (6)

ii. We induct on the assumption. There is one case.

Case (7.8). We have:

(1)	$\grave{r} = \mathtt{cerule}(p.\grave{e})$	by assumption
(2)	$r = \mathtt{rule}(p.e)$	by assumption
(3)	$\Delta \vdash p : \tau \dashv Y$	by assumption
(4)	$\Delta\,\Gamma \cup Y \vdash^{\hat{\Delta};\hat{\Gamma};\hat{\Psi};\hat{\Phi};b} \grave{e} \rightsquigarrow e \Rightarrow \tau'$	by assumption
(5)	$\Delta \cap \Delta_{\mathsf{app}} = \varnothing$	by assumption
(6)	$\mathsf{dom}(\Gamma) \cap \mathsf{dom}(Y) = \varnothing$	by identification convention
(7)	$\mathsf{dom}(\Gamma_{\mathsf{app}}) \cap \mathsf{dom}(Y) = \varnothing$	by identification convention
(8)	$\mathsf{dom}(\Gamma) \cap \mathsf{dom}(\Gamma_{\mathsf{app}}) = \varnothing$	by assumption
(9)	$\mathsf{dom}(\Gamma \cup Y) \cap \mathsf{dom}(\Gamma_{\mathsf{app}}) = \varnothing$	by standard finite set definitions and identities on (6), (7) and (8)
(10)	$\Delta \cup \Delta_{\mathsf{app}}\,\Gamma \cup Y \cup \Gamma_{\mathsf{app}} \vdash e : \tau'$	by IH, part 2(a)(i) on (4), (5) and (9)
(11)	$\Delta \cup \Delta_{\mathsf{app}}\,\Gamma \cup \Gamma_{\mathsf{app}} \cup Y \vdash e : \tau'$	by exchange of Y and Γ_{app} on (10)
(12)	$\Delta \cup \Delta_{\mathsf{app}}\,\Gamma \cup \Gamma_{\mathsf{app}} \vdash \mathtt{rule}(p.e) : \tau \Rrightarrow \tau'$	by Rule (4.2) on (3) and (11)

(b) i. We induct on the assumption.

Case (7.7a). We have:

(1)	$\Delta\,\Gamma \vdash^{\mathbb{E}} \grave{e} \rightsquigarrow e \Rightarrow \tau$	by assumption
(2)	$\Delta\,\Gamma \vdash e : \tau$	by IH, part 2(a)(i) on (1)

Case (7.7b) through (7.2h). In each of these cases, we apply:

- Lemma 3.19 to or over all ce-type validation premises.
- The IH, part 2(a)(i) to or over all synthetic ce-expression validation premises.
- The IH, part 2(a)(ii) to or over all synthetic ce-rule validation premises.
- The IH, part 2(b)(i) to or over all analytic ce-expression validation premises.

We then derive the conclusion by applying Rules (4.1), Rule (4.2), Lemma 4.1, the identification convention and exchange as needed.

Case (7.7i). We have:

(3) $\grave{e} = \mathtt{cesplicede}[m; n]$	by assumption
(4) $\mathsf{parseUExp}(\mathsf{subseq}(b; m; n)) = \hat{e}$	by assumption
(5) $\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \rightsquigarrow e \Leftarrow \tau$	by assumption
(6) $\Delta \cup \Delta_{\mathsf{app}} \vdash \tau\ \mathsf{type}$	by assumption
(7) $\Delta \cap \Delta_{\mathsf{app}} = \varnothing$	by assumption
(8) $\mathsf{dom}(\Gamma) \cap \mathsf{dom}(\Gamma_{\mathsf{app}}) = \varnothing$	by assumption
(9) $\Delta_{\mathsf{app}}\,\Gamma_{\mathsf{app}} \vdash e : \tau$	by IH, part 1(b)(i) on (5), (7), (8) and (6)
(10) $\Delta \cup \Delta_{\mathsf{app}}\,\Gamma \cup \Gamma_{\mathsf{app}} \vdash e : \tau$	by Lemma 4.1 over Δ and Γ and exchange on (9)

ii. We induct on the assumption. There is one case.

Case (7.9). We have:

(1) $\grave{r} = \mathtt{cerule}(p.\grave{e})$	by assumption
(2) $r = \mathtt{rule}(p.e)$	by assumption
(3) $\Delta \vdash p : \tau \dashv Y$	by assumption
(4) $\Delta\,\Gamma \cup Y \vdash^{\hat{\Delta};\hat{\Gamma};\hat{\Psi};\hat{\Phi};b} \grave{e} \rightsquigarrow e \Leftarrow \tau'$	by assumption
(5) $\Delta \cup \Delta_{\mathsf{app}} \vdash \tau'\ \mathsf{type}$	by assumption
(6) $\mathsf{dom}(\Gamma) \cap \mathsf{dom}(\Gamma_{\mathsf{app}}) = \varnothing$	by assumption
(7) $\Delta \cap \Delta_{\mathsf{app}} = \varnothing$	by assumption
(8) $\mathsf{dom}(\Gamma) \cap \mathsf{dom}(Y) = \varnothing$	by identification convention
(9) $\mathsf{dom}(\Gamma_{\mathsf{app}}) \cap \mathsf{dom}(Y) = \varnothing$	by identification convention
(10) $\mathsf{dom}(\Gamma \cup Y) \cap \mathsf{dom}(\Gamma_{\mathsf{app}}) = \varnothing$	by standard finite set definitions and identities on (8), (9) and (6)
(11) $\Delta \cup \Delta_{\mathsf{app}}\,\Gamma \cup Y \cup \Gamma_{\mathsf{app}} \vdash e : \tau'$	by IH, part 2(b)(i) on (4), (7), (10) and (5)
(12) $\Delta \cup \Delta_{\mathsf{app}}\,\Gamma \cup \Gamma_{\mathsf{app}} \cup Y \vdash e : \tau'$	by exchange of Y and Γ_{app} on (11)

$$(13) \quad \Delta \cup \Delta_{\mathrm{app}} \ \Gamma \cup \Gamma_{\mathrm{app}} \vdash \mathtt{rule}(p.e) : \tau \mapsto \tau' \text{ by Rule (4.2) on (3)}$$
$$\text{and (12)}$$

We must now show that the induction is well-founded. All applications of the IH are on subterms except the following.

- The only cases in the proof of part 1 that invoke the IH, part 2 are Case (7.1m) in the proof of part 1(a)(i) and Case (7.2k) in the proof of part 1(b)(i). The only cases in the proof of part 2 that invoke the IH, part 1 are Case (7.6l) in the proof of part 2(a)(i) and Case (7.7i) in the proof of part 2(b)(i). We can show that the following metric on the judgements that we induct on is stable in one direction and strictly decreasing in the other direction:

$$\|\hat{\Delta}\ \hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \rightsquigarrow e \Rightarrow \tau\| = \|\hat{e}\|$$
$$\|\hat{\Delta}\ \hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \rightsquigarrow e \Leftarrow \tau\| = \|\hat{e}\|$$
$$\|\Delta\ \Gamma \vdash^{\hat{\Delta};\hat{\Gamma};\hat{\Psi};\hat{\Phi};b} \hat{e} \rightsquigarrow e \Rightarrow \tau\| = \|b\|$$
$$\|\Delta\ \Gamma \vdash^{\hat{\Delta};\hat{\Gamma};\hat{\Psi};\hat{\Phi};b} \hat{e} \rightsquigarrow e \Leftarrow \tau\| = \|b\|$$

where $\|b\|$ is the length of b and $\|\hat{e}\|$ is the sum of the lengths of the ueTSM literal bodies in $\hat{e}$,

$$\|\hat{x}\| = 0$$
$$\|\mathtt{uasc}\{\hat{\tau}\}(\hat{e})\| = \|\hat{e}\|$$
$$\|\mathtt{uletval}(\hat{e}; \hat{x}.\hat{e}')\| = \|\hat{e}\| + \|\hat{e}'\|$$
$$\|\mathtt{uanalam}(\hat{x}.\hat{e})\| = \|\hat{e}\|$$
$$\|\mathtt{ulam}\{\hat{\tau}\}(\hat{x}.\hat{e})\| = \|\hat{e}\|$$
$$\|\mathtt{uap}(\hat{e}_1; \hat{e}_2)\| = \|\hat{e}_1\| + \|\hat{e}_2\|$$
$$\|\mathtt{utlam}(\hat{t}.\hat{e})\| = \|\hat{e}\|$$
$$\|\mathtt{utap}\{\hat{\tau}\}(\hat{e})\| = \|\hat{e}\|$$
$$\|\mathtt{ufold}(\hat{e})\| = \|\hat{e}\|$$
$$\|\mathtt{uunfold}(\hat{e})\| = \|\hat{e}\|$$
$$\|\mathtt{utpl}[L](\{i \hookrightarrow \hat{e}_i\}_{i \in L})\| = \sum_{i \in L} \|\hat{e}_i\|$$
$$\|\mathtt{upr}[\ell](\hat{e})\| = \|\hat{e}\|$$
$$\|\mathtt{uin}[\ell](\hat{e})\| = \|\hat{e}\|$$
$$\|\mathtt{umatch}[n](\hat{e}; \{\hat{r}_i\}_{1 \le i \le n})\| = \|\hat{e}\| + \sum_{1 \le i \le n} \|r_i\|$$
$$\|\mathtt{usyntaxue}\{e_{\mathrm{parse}}\}\{\hat{\tau}\}(\hat{a}.\hat{e})\| = \|\hat{e}\|$$
$$\|\mathtt{uimplicite}[\hat{a}](\hat{e})\| = \|\hat{e}\|$$
$$\|\mathtt{uapuetsm}[b][\hat{a}]\| = \|b\|$$

$$\|\texttt{uelit}[b]\| = \|b\|$$
$$\|\texttt{usyntaxup}\{e_{\text{parse}}\}\{\hat{\tau}\}(\hat{a}.\hat{e})\| = \|\hat{e}\|$$
$$\|\texttt{uimplicitp}[\hat{a}](\hat{e})\| = \|\hat{e}\|$$

and $\|r\|$ is defined as follows:

$$\|\texttt{urule}(\hat{p}.\hat{e})\| = \|\hat{e}\|$$

Going from part 1 to part 2, the metric remains stable:

$$\|\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \texttt{uapuetsm}[b][\hat{a}] \rightsquigarrow e \Rightarrow \tau\|$$
$$= \|\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \texttt{uelit}[b] \rightsquigarrow e \Leftarrow \tau\|$$
$$= \|\varnothing\,\varnothing \vdash^{\hat{\Delta};\hat{\Gamma};\hat{\Psi};\hat{\Phi};b} \grave{e} \rightsquigarrow e \Leftarrow \tau\|$$
$$= \|b\|$$

Going from part 2 to part 1, in each case we have that $\mathsf{parseUExp}(\mathsf{subseq}(b; m; n)) = \hat{e}$ and the IH is applied to the judgements $\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \rightsquigarrow e \Rightarrow \tau$ and $\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \rightsquigarrow e \Leftarrow \tau$, respectively. Because the metric is stable when passing from part 1 to part 2, we must have that it is strictly decreasing in the other direction:

$$\|\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \rightsquigarrow e \Rightarrow \tau\| < \|\Delta\,\Gamma \vdash^{\hat{\Delta};\hat{\Gamma};\hat{\Psi};\hat{\Phi};b} \texttt{cesplicede}[m;n] \rightsquigarrow e \Rightarrow \tau\|$$

and

$$\|\hat{\Delta}\,\hat{\Gamma} \vdash_{\hat{\Psi};\hat{\Phi}} \hat{e} \rightsquigarrow e \Leftarrow \tau\| < \|\Delta\,\Gamma \vdash^{\hat{\Delta};\hat{\Gamma};\hat{\Psi};\hat{\Phi};b} \texttt{cesplicede}[m;n] \rightsquigarrow e \Leftarrow \tau\|$$

i.e. by the definitions above,

$$\|\hat{e}\| < \|b\|$$

This is established by appeal to Condition 3.22, which states that subsequences of b are no longer than b, and the following condition, which states that an unexpanded expression constructed by parsing a textual sequence b is strictly smaller, as measured by the metric defined above, than the length of b, because some characters must necessarily be used to delimit each literal body.

Condition 7.9 (Expression Parsing Monotonicity). *If* $\mathsf{parseUExp}(b) = \hat{e}$ *then* $\|\hat{e}\| < \|b\|$.

Combining Conditions 3.22 and 7.9, we have that $\|\hat{e}\| < \|b\|$ as needed.

- In Case (7.2a) of the proof of part 1(b)(i), we apply the IH, part 1(a)(i), with $\hat{e} = \hat{e}$. This is well-founded because all applications of the IH, part 1(b)(i) elsewhere in the proof are on strictly smaller terms.
- Similarly, in Case (7.7a) of the proof of part 2(b)(i), we apply the IH, part 2(a)(i), with $\grave{e} = \grave{e}$. This is well-founded because all applications of the IH, part 2(b)(i) elsewhere in the proof are on strictly smaller terms.

$\square$

7.3 Related Work

TODO: cite/comment on past work on bidirectional typechecking TODO: TSLs TODO: ichikawa modularity paper TODO: other things from related work section of ECOOP paper

Part IV

Conclusion

Chapter 8

Discussion & Future Directions

8.1 Interesting Applications

Most of the examples in Sec. 2.3 can be expressed straightforwardly using the constructs introduced in the previous chapters. Here, let us highlight certain interesting examples and exceptions.

8.1.1 Monadic Commands

8.2 Summary

TODO: Write summary

8.3 Future Directions

We did not consider situations where a library clients wants to provide derived forms for defining types, signatures, modules or other declarations (though we have explored syntax for types in a recent short paper [51].) We also did not consider situations where a library client wants to generate an expression or a pattern based on the structure of a type (e.g. automatic generation of equality comparisons.) Finally, we do not consider situations that require modifications to the underlying type structure of a language (though "reasonably programmable type structure" is a rich avenue for future work.)

8.3.1 TSM Packaging

In the exposition thusfar, we have assumed that TSMs have delimited scope. However, ideally, we would like to be able to define TSMs within a module:

```
structure Rxlib = struct
  type Rx = (* ... *)
  syntax $rx at Rx { (* ... *) }
```

```
  (* ... *)
  end
```

However, this leads to an important question: how can we write down a signature for the module `Rxlib`? One approach would be to simply duplicate the full definition of the TSM in the signature, but this leads to inelegant code duplication and raises the difficult question of how the language should decide whether one TSM is a duplicate of another. For this reason, in VerseML, a signature can only refer to a previously defined TSM. So, for example, we can write down a signature for `Rxlib` after it has been defined:

```
signature RXLIB = sig
  type Rx = (* ... *)
  syntax $rx = Rxlib.$rx
  (* ... *)
end
Rxlib : RXLIB (* check Rxlib against RXLIB after the fact *)
```

Alternatively, we can define the type `Rx` and the TSM `$rx` before defining `Rxlib`:

```
local
  type Rx = (* ... *)
  syntax $rx at Rx { (* ... *) }
in
  structure Rxlib :
  sig
    type Rx = Rx
    syntax $rx = $rx
    (* ... *)
  end = struct
    type Rx = Rx
    syntax $rx = $rx
    (* ... *)
  end
end
```

Another important question is: how does a TSM defined within a module at a type that is held abstract outside of that module operate? For example, consider the following:

```
local
  type Rx = (* ... *)
  syntax $rx at Rx { (* ... *) }
in
  structure Rxlib :
  sig
    type Rx (* held abstract *)
    syntax $rx = $rx
    (* ... *)
  end = struct
    type Rx = Rx
    syntax $rx = $rx
    (* ... *)
  end
```

```
    end
```

If we apply `Rxlib.$rx`, it may generate an expansion that uses the constructors of the `Rx` type. However, because the type is being held abstract, these constructors may not be visible at the application site. TODO: actually, this is why doing this is a bad idea. export TSMs only from units, not modules

8.3.2 TSLs

8.4 pTSLs By Example

For example, a module `P` can associate the TSM `rx` defined in the previous section with the abstract type `R.t` by qualifying the definition of the sealed module it is defined by as follows:

```
module R = mod {
  type t = (* ... *)
  (* ... *)
} :> RX with syntax rx
```

More generally, when sealing a module expression against a signature, the programmer can specify, for each abstract type that is generated, at most one previously defined TSMs. This TSM must take as its first parameter the module being sealed.

The following function has the same expansion as `example_using_tsm` but, by using the TSL just defined, it is more concise. Notice the return type annotation, which is necessary to ensure that the TSL can be unambiguously determined:

```
fun example_using_tsl(name : string) : R.t => /@name: %ssn/
```

As another example, let us consider the standard list datatype. We can use TSLs to express derived list syntax, for both expressions and patterns:

```
datatype list('a) { Nil | Cons of 'a * list('a) } with syntax {
  static fn (body : Body) =>
    (* ... comma-delimited spliced exps ... *)
} with pattern syntax {
  static fn (body : Body) : Pat =>
    (* ... list pattern parser ... *)
}
```

Together with the TSL for regular expression patterns, this allows us to write lists like this:

```
let val x : list(R.t) = [/\d/, /\d\d/, /\d\d\d/]
```

From the client's perspective, it is essentially as if the language had built in derived syntax for lists and regular expression patterns directly.

8.5 Parameterized Modules

TSLs can be associated with abstract types that are generated by parameterized modules (i.e. generative functors in Standard ML) as well. For example, consider a trivially parameterized module that creates modules sealed against RX:

```
module F() => mod {
  type t = (* ... *)
  (* ... *)
} :> RX with syntax rx
```

Each application of F generates a distinct abstract type. The semantics associates the appropriately parameterized TSM with each of these as they are generated:

```
module F1 = F() (* F1.t has TSL rx(F1) *)
module F2 = F() (* F2.t has TSL rx(F2) *)
```

As a more complex example, let us define two signatures, A and B, a TSM $G and a parameterized module G : A -> B:

```
signature A = sig { type t; val x : t }
signature B = sig { type u; val y : u }
syntax $G(M : A)(G : B) at G.u { (* ... *) }
module G(M : A) => mod {
  type u = M.t; val y = M.x } :> B with syntax $G(M)
```

Both G and $G take a parameter M : A. We associate the partially applied TSM $G(M) with the abstract type that G generates. Again, this satisfies the requirement that one must be able to apply the TSM being associated with the abstract type to the module being sealed.

Only fully abstract types can have TSLs associated with them. Within the definition of G, type u does not have a TSL available to it because it is synonymous to M.t. More generally, TSL lookup respects type equality, so any synonyms of a type with a TSL will also have that TSL. We can see this in the following example, where the type u has a different TSL associated with it inside and outside the definition of the module N:

```
module M : A = mod { type t = int; val x = 0 }
module G1 = G(M) (* G1.t has TSL $G(M), per above *)
module N = mod {
  type u = G1.t (* u = G1.t in this scope, so u also has TSL $G(M) *)
  val y = /asdf/ (* we can use it to create a value of that type *)
} :> B (* did not specify a TSL for N.u at the point where it is sealed,
          so N.u has no TSL in the outer scope *)
val z : N.u = /asdf/ (* ERROR: no TSL for type N.u *)
```

8.6 miniVerse_TSL

A formal specification of TSLs in a language that supports only non-parametric datatypes is available in a paper published in ECOOP 2014 [50].

8.6.1 TSMs and TSLs In Candidate Expansions

Candidate expansions cannot themselves define or apply TSMs. This simplifies our metatheory, though it can be inconvenient at times for TSM providers. We discuss adding the ability to use TSMs within candidate expansions here.TODO: write this

8.6.2 Pattern Matching Over Values of Abstract Type

ML does not presently support pattern matching over values of an abstract data type. However, there have been proposals for adding support for pattern matching over abstract data types defined by modules having a "datatype-like" shape, e.g. those that define a case analysis function like the one specified by RX,

8.6.3 Integration Into Other Languages

We conjecture that the constructs we describe could be integrated into dependently typed functional languages, e.g. Coq, but leave the technical developments necessary for doing so as future work.

Some of the constructs in Chapter 3, Chapter 5 and Chapter 7 could also be adapted for use in imperative languages with non-trivial type structure, like Java. Similarly, some of the constructs we discuss could also be adapted into "dynamic languages" like Racket or Python, though the constructs in Chapter 7 are not relevant to such languages.

8.6.4 Mechanically Verifying TSM Definitions

Finally, VerseML is not designed for advanced theorem proving tasks where languages like Coq, Agda or Idris might be used today. That said, we conjecture that the primitives we describe could be integrated into languages like Gallina (the "external language" of the Coq proof assistant [45]) with modifications, but do not plan to pursue this line of research here.

In such a setting, you could verify TSM definitions TODO: finish writing this

8.6.5 Improved Error Reporting

8.6.6 Controlled Binding

8.6.7 Type-Aware Splicing

8.6.8 Integration With Code Editors

8.6.9 Resugaring

TODO: Cite recent work at PLDI (?) and ICFP from Brown

8.6.10 Non-Textual Display Forms

TODO: Talk about active code completion work and future ideas

LATEX Source Code and Updates

The LATEX sources for this document can be found at the following URL:

<div align="center">

https://github.com/cyrus-/thesis

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Any errors or omissions can be reported to the author by email at the following address:

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</div>

The author will also review and accept pull requests on GitHub.

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TODO (Later): Remove extraneous nonsense from entries.

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