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PAPER

Controlling unfolding in type theory

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Abstract

We present a new way to control the unfolding of definitions in dependent type theory. Traditionally, proof assistants require users to fix whether each definition will or will not be unfolded in the remainder of a development; unfolding definitions is often necessary in order to reason about them, but an excess of unfolding can result in brittle proofs and intractably large proof goals. In our system, definitions are by default not unfolded, but users can selectively unfold them in a local manner. We justify our mechanism by means of elaboration to a core theory with *extension types* – a connective first introduced in the context of homotopy type theory – and by establishing a normalization theorem for our core calculus. We have implemented controlled unfolding in the *cooltt* proof assistant, inspiring an independent implementation in Agda.

Keywords: dependent type theory; proof assistants; categorical type theory

1. Introduction

In dependent type theory, terms are type-checked modulo definitional equality, a congruence generated by α -, β -, and η -laws, as well as unfolding of definitions. Unfolding definitions is to some extent a convenience that allows type checkers to silently discharge many proof obligations, for example, a list of length $1 + 1$ is without further annotation also a list of length 2. It is by no means the case, however, that we always want a given definition to unfold:

- *Modularity:* Dependent types are famously sensitive to the smallest changes to definitions, such as whether $(+)$ recurs on its first or its second argument. If we plan to change a definition in the future, it may be desirable to avoid exposing its implementation to the type checker.
- *Usability:* While unfolding may simplify proof states, it also has the potential to complicate them, resulting in unreadable subgoals, error messages, *etc.* A user may find that certain definitions are likely to be problematic in this way and thus opt not to unfold them.

Many proof assistants accordingly have implementation-level support for marking definitions *opaque* (unable to be unfolded), including Agda’s `abstract` (The Agda Team 2021) and Coq’s `Qed` (The Coq Development Team 2022). But unfolding definitions is not merely a matter of convenience: to reason about a function, we must unfold it. For example, if we make the definition of

(+) opaque, then (+) is indistinguishable from a variable of type $\mathbb{N} \rightarrow \mathbb{N} \rightarrow \mathbb{N}$ and so cannot be shown to be commutative, satisfy $1 + 1 = 2$, etc.

In practice, proof assistants resolve this contradiction by adopting an intermediate stance: definitions are *transparent* (unfolded during type checking) by default, but users are given some control over their unfolding. Coq provides conversion tactics (`cbv`, `simpl`, etc.) for applying definitional equalities, each of which accepts a list of definitions to unfold; its `Opaque` and `Transparent` commands toggle the default unfolding behavior of a transparent definition; and the `SSReflect` tactic language natively supports a “locking” idiom for controlling when definitions unfold (Gonthier et al. 2016). Agda allows users to group multiple definitions into a single `abstract` block, inside of which those definitions are transparent and outside of which they are opaque; this allows users to define a function, prove all lemmas that depend on the function’s definition, and then irreversibly make the function and lemmas opaque.

These mechanisms for controlling unfolding pose interesting trade-offs for users: which definitions should be transparent, and which should be opaque? Transparency is in some cases necessary and in many cases convenient, but it is problematic both from an engineering perspective – because any edit to a transparent definition can break the well-typedness of any number of its use sites – and from a performance perspective – because checking definitional equality of type indices often requires unfolding nested definitions into large normal forms.

In addition, the behavior of these mechanisms is more subtle than it may at first appear. In Agda, definitions within `abstract` blocks are transparent to other definitions in the same block, but opaque to the *types* of those definitions; without such a stipulation, those types may cease to be well-formed when the earlier definition is made opaque. Furthermore, `abstract` blocks are anti-modular, requiring users to anticipate all future lemmas about definitions in a block.¹ Coq’s conversion tactics are more flexible than Agda’s `abstract` blocks, but being tactics, their behavior can be harder to predict. The `lock` idiom in `SSReflect` is more predictable because it creates opaque definitions but comes in four different variations to simplify its use in practice.

1.1 Contributions

We propose a mechanism for fine-grained control over the unfolding of definitions in dependent type theory. We introduce language-level primitives for *controlled unfolding* that are elaborated into a core calculus with *extension types*, a connective first introduced by Riehl and Shulman (2017). We justify our elaboration algorithm by establishing a normalization theorem (and hence the decidability of type checking and injectivity of type constructors) for our core calculus, and we have implemented our system for controlled unfolding in the experimental `cooltt` proof assistant (RedPRL Development Team 2020).

Definitions in our framework are opaque by default but can be selectively and locally unfolded as if they were transparent. Our system is finer-grained and more modular than Agda’s `abstract` blocks: we need not collect all lemmas that unfold a given definition into a single block, making our mechanism better suited to libraries. Our primitives have more predictable meaning and performance than Coq’s unfolding tactics² because they are implemented by straightforward elaboration into a core Martin-Löf type theory (MLTT) extended with new types and declaration forms.

In particular, we refine earlier approaches to representing definitions within type theory (Dreyer et al. 2003; Harper & Stone 2000; Milner et al. 1997; Sterling and Harper 2021) in order to more faithfully represent definitions as they are actually used in practice: as neither fully opaque or transparent but instead a mix of the two. Drawing inspiration from cubical type theory (Angiuli et al. 2018, 2021; Cohen et al. 2017), we extend MLTT with proof-irrelevant proposition symbols p , dependent products $\{p\} A$ over those propositions, and extension types $\{A|p \hookrightarrow a\}$, the subtype of A consisting of the elements of A that definitionally equal a under the assumption that p is true. For readers familiar with cubical type theory, extension types are similar to path types

(Path $A \ a_0 \ a_1$), which classify functions out of an abstract interval \mathbb{I} that are definitionally equal to a_0 and a_1 when evaluated at the interval's endpoints $0, 1 : \mathbb{I}$.

Encoding definitions through particular types confers a number of benefits. For instance, our mechanism for definitions and unfolding are automatically invariant under definitional equivalence: replacing one term by a definitionally equal alternative cannot change the unfolding behavior of a program. Furthermore, using extension types to encode definitions ensures that our elaboration algorithm is extremely modular and predictable: the rules for extension types are simple and, once grasped, it becomes easy to predict the interactions between unfolding definitions and other features within the language. This elaboration algorithm then serves as a *reference* for the behavior of our mechanism, against which other implementation strategies may be checked.

Like many elaboration algorithms for dependent type theory, executing our elaboration algorithm requires deciding the equality of types in the core language. To show that our elaboration algorithm can be implemented, we prove a *normalization* theorem for our core calculus, characterizing its definitional equivalence classes of types and terms and as a corollary establishing the decidability of type checking. This is more subtle than it may appear: the heart of our normalization proof amounts to correctly tracking when definitions are allowed to unfold as well as when they should remain opaque. In the face of higher-order programs and dependent types, this is quite difficult.

Another benefit to shifting from opaque definitions to extension types is their well-studied metatheory. Specifically, we are able to adapt and extend Sterling's technique of synthetic Tait computability (STC) (Sterling 2021; Sterling and Angiuli 2021; Sterling and Harper 2021) to prove normalization for our core language. Our proof is fully constructive, an improvement on the prior work of Sterling and Angiuli (2021); we have also corrected an error in the handling of universes in an earlier revision of Sterling's doctoral dissertation (Sterling 2021) that was detected while preparing this paper.

1.2 Outline

In Section 2, we introduce our controlled unfolding primitives by way of examples, and in Section 3 we walk through how these examples are elaborated into our core language of type theory with proposition symbols and extension types. In Section 4, we present our elaboration algorithm, and in Section 5 we discuss our implementation of the above in the *cooltt* proof assistant. In Section 6, we establish normalization and its corollaries for our core calculus. We conclude with a discussion of related work in Section 7.

2. A Surface Language with Controlled Unfolding

We begin by describing an Agda-like surface language for a dependent type theory with controlled unfolding. In Section 4, we will give precise meaning to this language by explaining how to elaborate it into our core calculus; for now, we proceed by example, introducing our new primitives bit by bit. Our examples will concern the inductively defined natural numbers and their addition function:

$$(+): \mathbb{N} \rightarrow \mathbb{N} \rightarrow \mathbb{N}$$

$$\text{ze} + n = n$$

$$\text{su } m + n = \text{su } (m + n)$$

2.1 A simple dependency: length-indexed vectors

In our language, definitions such as $(+)$ are opaque by default – they are not unfolded automatically. To illustrate the need to *selectively* unfold $(+)$, consider the indexed inductive type of length-indexed vectors with the following constructors:

$$[] : \text{vec } \text{ze } A$$

$$(::) : A \rightarrow \text{vec } n A \rightarrow \text{vec } (\text{su } n) A$$

Suppose we attempt to define the *append* operation on vectors by dependent pattern matching on the first vector. Our goals would be as follows:

$$(\oplus) : \text{vec } m A \rightarrow \text{vec } n A \rightarrow \text{vec } (m + n) A$$

$$[] \oplus v = ? : \text{vec } (\text{ze} + n) A$$

$$(a :: u) \oplus v = ? : \text{vec } (\text{su } m + n) A$$

As it stands, the goals above are in normal form and cannot be proved; however, we may indicate that the definition of $(+)$ should be unfolded within the definition of (\oplus) by adding the following top-level **unfolds** annotation:

(\oplus) unfolds $(+)$

$$(\oplus) : \text{vec } m A \rightarrow \text{vec } n A \rightarrow \text{vec } (m + n) A$$

With our new declaration, the goals simplify:

$$[] \oplus v = ? : \text{vec } n A$$

$$(a :: u) \oplus v = ? : \text{vec } (\text{su } (m+n)) A$$

The first goal is solved with v itself; for the second goal, we begin by applying the `vcons` constructor:

$$(a :: u) \oplus v = a :: ? : \text{vec } (m+n) A$$

The remaining goal is just our induction hypothesis $u \oplus v$. All in all, we have:

(\oplus) unfolds $(+)$

$$(\oplus) : \text{vec } m A \rightarrow \text{vec } n A \rightarrow \text{vec } (m + n) A$$

$$[] \oplus v = v$$

$$(a :: u) \oplus v = a :: (u \oplus v)$$

2.2 Transitive unfolding

Now suppose we want to prove that `map` distributes over (\oplus) . In doing so, we will certainly need to unfold `map`, but it turns out this will not be enough:

$$\text{map} : (A \rightarrow B) \rightarrow \text{vec } n A \rightarrow \text{vec } n B$$

$$\text{map } f [] = []$$

$$\text{map } f (a :: u) = f a :: \text{map } f u$$

map- \oplus unfolds map

$\text{map-}\oplus : (f : A \rightarrow B) (u : \text{vec } m A) (v : \text{vec } n A) \rightarrow \text{map } f (u \oplus v) \equiv \text{map } f u \oplus \text{map } f v$

$\text{map-}\oplus f [] v = ? : \text{map } f ([] \oplus v) \equiv [] \oplus \text{map } f v$

$\text{map-}\oplus f (a :: u) v = ? : \text{map } f (a :: u) \oplus v \equiv (f a :: \text{map } f u) \oplus \text{map } f v$

To make further progress, we must also unfold (\oplus):

map- \oplus unfolds map;(\oplus)

$\text{map-}\oplus : (f : A \rightarrow B) (u : \text{vec } m A) (v : \text{vec } n A) \rightarrow \text{map } f (u \oplus v) \equiv \text{map } f u \oplus \text{map } f v$

$\text{map-}\oplus f [] v = ? : \text{map } f v \equiv \text{map } f v$

$\text{map-}\oplus f (a :: u) v = ? : f a :: \text{map } f (u \oplus v) \equiv (f a :: \text{map } f u) \oplus \text{map } f v$

In our language, unfolding (\oplus) has the side effect of *also* unfolding ($+$): in other words, unfolding is *transitive*. To see why this is the case, observe that the unfolding of $(a :: u) \oplus v : \text{vec } (\text{su } m + n) A$, namely $a :: (u \oplus v) : \text{vec } (\text{su } (m + n)) A$, would otherwise not be well-typed. From an implementation perspective, one can think of the transitivity of unfolding as necessary for *subject reduction*. Having unfolded map , (\oplus), and thus ($+$), we complete our definition:

$\text{cong} : (f : A \rightarrow B) \rightarrow a \equiv a' \rightarrow f a \equiv f a'$

$\text{cong } f \text{ refl} = \text{refl}$

map- \oplus unfolds map;(\oplus)

$\text{map-}\oplus : (f : A \rightarrow B) (u : \text{vec } m A) (v : \text{vec } n A) \rightarrow \text{map } f (u \oplus v) \equiv \text{map } f u \oplus \text{map } f v$

$\text{map-}\oplus f [] v = \text{refl}$

$\text{map-}\oplus f (a :: u) v = \text{cong } (f a ::) (\text{map-}\oplus f u v)$

2.3 Recovering unconditionally transparent/opaque definitions

There are also times when we intend a given definition to be a fully transparent *abbreviation*, in the sense of being unfolded automatically whenever possible. We indicate this with an **abbreviation** declaration:

abbreviation singleton

$\text{singleton} : A \rightarrow \text{vec } (\text{su } ze) A$

$\text{singleton } a = a :: []$

Then the following lemma can be defined without any explicit unfolding:

$\text{abbrv-example} : \text{singleton } 5 \equiv (5 :: [])$

$\text{abbrv-example} = \text{refl}$

The meaning of the **abbreviation** keyword must account for unfolding constraints. For instance, what would it mean to make $\text{map-}\oplus$ an abbreviation?

abbreviation map- \oplus
 map- \oplus **unfolds** map; (\oplus)
 ...

We cannot unfold map- \oplus in all contexts, because its definition is only well-typed when map and (\oplus) are unfolded. The meaning of this declaration must, therefore, be that map- \oplus shall be unfolded *just as soon as* map and (\oplus) are unfolded. In other words, **abbreviation** ϑ followed by ϑ **unfolds** $\kappa_1; \dots; \kappa_n$ means that unfolding ϑ is synonymous with unfolding all of $\kappa_1; \dots; \kappa_n$.

Conversely, we may intend a given definition *never* to unfold, which we may indicate by a corresponding **abstract** declaration. Because definitions in our system do not automatically unfold, the force of **abstract** ϑ is simply to prohibit users from including ϑ in any subsequent **unfolds** annotations.

Remark 1. A slight variation on our system can recover the behavior of Agda’s **abstract** blocks by limiting the scope in which a definition ϑ can be unfolded; the transitivity of unfolding dictates that any definition ϑ' that unfolds ϑ cannot itself be unfolded once we leave that scope. We leave the details to future work.

2.4 Unfolding within the type

The effect of a ϑ **unfolds** $\kappa_1; \dots; \kappa_n$ declaration is to make $\kappa_1; \dots; \kappa_n$ unfold within the *definition* of ϑ , but still not within its type; it will happen, however, that a *type* might not be expressible without some unfolding. First, we will show how to accommodate this situation using only features we have introduced so far, and then in Section 2.5, we will devise a more general and ergonomic solution.

Consider the left-unit law for (\oplus): in order to state that a vector u is equal to the vector $[] \oplus u$, we must contend with their differing types $\text{vec } n \ A$ and $\text{vec } (ze + n) \ A$, respectively. One approach is to rewrite along the left-unit law for \mathbb{N} ; indeed, to state the *right-unit* law for (\oplus), one must rewrite along the right-unit law for \mathbb{N} . But here, because $(+)$ computes on its first argument, $\text{vec } n \ A$ and $\text{vec } (ze + n) \ A$ would be definitionally equal types if we could unfold $(+)$.

In order to formulate the left-unit law for (\oplus), we start by defining its *type* as an abbreviation that unfolds $(+)$:

abbreviation \oplus -left-unit-type
 \oplus -left-unit-type **unfolds** $(+)$
 \oplus -left-unit-type : $\text{vec } n \ A \rightarrow \text{Type}$
 \oplus -left-unit-type $u = [] \oplus u \equiv u$

Now we may state the intended lemma using the type defined above:

\oplus -left-unit : $(u : \text{vec } n \ A) \rightarrow \oplus$ -left-unit-type u
 \oplus -left-unit $u = ? : \oplus$ -left-unit-type u

Clearly, we must unfold $(+)$ and thus \oplus -left-unit-type to simplify our goal:

\oplus -left-unit **unfolds** $(+)$
 \oplus -left-unit : $(u : \text{vec } n \ A) \rightarrow \oplus$ -left-unit-type u
 \oplus -left-unit $u = ? : [] \oplus u \equiv u$

We complete the proof by unfolding (\oplus) itself, which transitively unfolds $(+)$:

\oplus -left-unit **unfolds** (\oplus)
 \oplus -left-unit : $(u : \text{vec } n A) \rightarrow \oplus$ -left-unit-type u
 \oplus -left-unit $u = \text{refl}$

2.5 Unfolding within subexpressions

We have just demonstrated how to unfold definitions within the *type* of a declaration by defining that type as an additional declaration; using the same technique, we can introduce unfoldings within *any subexpression* by hoisting that subexpression to a top-level definition with its own unfolding constraint.

Unfolding within the type, revisited. Rather than repeating the somewhat verbose pattern of Section 2.4, we abstract it as a new language feature that is easily eliminated by elaboration. In particular, we introduce a new *expression* former **unfold** κ **in** M that can be placed in any expression context. Let us replay the example from Section 2.4, but using **unfold** rather than an auxiliary definition:

\oplus -left-unit : $(u : \text{vec } n A) \rightarrow$ **unfold** $(+)$ **in** $[] \oplus u \equiv u$
 \oplus -left-unit $u =$ **?** : **unfold** $(+)$ **in** $[] \oplus u \equiv u$

The type **unfold** $(+)$ **in** $[] \oplus u \equiv u$ is in normal form; the only way to simplify it is to unfold $(+)$. We could do this with another inline **unfold** expression (see \oplus -left-unit' below), but here we will use a top-level declaration:

\oplus -left-unit **unfolds** $(+)$
 \oplus -left-unit : $(u : \text{vec } n A) \rightarrow$ **unfold** $(+)$ **in** $[] \oplus u \equiv u$
 \oplus -left-unit $u =$ **?** : $[] \oplus u \equiv u$

By virtue of the above, the **unfold** expression in our hole has computed away and we are left with **?** : $[] \oplus u \equiv u$ as (\oplus) is still abstract in this scope. To make progress, we *strengthen* the declaration to unfold (\oplus) in addition to $(+)$:

\oplus -left-unit **unfolds** (\oplus)
 \oplus -left-unit : $(u : \text{vec } n A) \rightarrow$ **unfold** $(+)$ **in** $[] \oplus u \equiv u$
 \oplus -left-unit $u = \text{refl}$

The meaning of the code above is exactly as described in Section 2.4: the **unfold** scope is elaborated to a new top-level **abbreviation** that unfolds $(+)$.

Expression-level vs. top-level unfolding. We noted in our definition of \oplus -left-unit above that we could have replaced the top-level **unfolds** (\oplus) directive of \oplus -left-unit with the new expression-level **unfold** (\oplus) **in** as follows:

\oplus -left-unit' : $(u : \text{vec } n A) \rightarrow$ **unfold** $(+)$ **in** $[] \oplus u \equiv u$
 \oplus -left-unit' $u =$ **unfold** (\oplus) **in** refl

The resulting definition of \oplus -left-unit' has slightly different behavior than \oplus -left-unit above: whereas unfolding \oplus -left-unit causes (\oplus) to unfold transitively, we can unfold \oplus -left-unit'

without unfolding (\oplus) – at the cost of **unfold** (\oplus) expressions appearing in our goal. This more granular behavior may be desirable in some cases, and it is a strength of our language and its elaborative semantics that the programmer can manipulate unfolding in such a fine-grained manner.

For completeness, we show the elaborated version of \oplus -left-unit' resulting from eliminating expression-level unfolding from the definition. We defer a systematic discussion of this transformation till Section 4.

abbreviation \oplus -left-unit'-type

\oplus -left-unit'-type **unfolds** (+)

\oplus -left-unit'-type : $\text{vec } n \ A \rightarrow \text{Type}$

\oplus -left-unit'-type $u = [] \oplus u \equiv u$

abbreviation \oplus -left-unit'-body

\oplus -left-unit'-body **unfolds** (\oplus)

\oplus -left-unit'-body : $(u : \text{vec } n \ A) \rightarrow \oplus\text{-left-unit'-type } u$

\oplus -left-unit'-body $u = \text{refl}$

$\oplus\text{-left-unit}' : (u : \text{vec } n \ A) \rightarrow \oplus\text{-left-unit'-type } u$

$\oplus\text{-left-unit}' \ u = \oplus\text{-left-unit'-body } u$

In our experience, expression-level unfolding seems more commonly useful for end users than top-level unfolding; on the other hand, the clearest semantics for expression-level unfolding are stated in terms of top-level unfolding. Because one of our goals is to provide an account of unfolding that admits a reliable and precise mental model for programmers, it is desirable to include both top-level and expression-level unfolding in the surface language.

3. Controlling Unfolding with Extension Types

Having introduced our new surface language constructs for controlled unfolding in Section 2, we now describe how to elaborate these constructs into our dependently typed core calculus. Again we proceed by example, deferring our formal descriptions of the elaboration algorithm to Section 4.

3.1 A core calculus with proposition symbols

Our core calculus parameterizes intensional MLTT (Martin-Löf 1975) by a bounded meet semilattice of *proposition symbols* $p \in \mathbb{P}$ and adjoins to the type theory a new form of context extension and two new type formers $\{p\} A$ and $\{A\} p \hookrightarrow M$ involving proposition symbols:

$$\begin{aligned} (\text{contexts}) \quad \Gamma &::= \dots \mid \Gamma, p \\ (\text{types}) \quad A &::= \dots \mid \{p\} A \mid \{A\} p \hookrightarrow M \end{aligned}$$

The bounded meet semilattice structure on \mathbb{P} closes proposition symbols under conjunction \wedge and the true proposition \top , thereby partially ordering \mathbb{P} by entailment $p \leq q$ (“ p entails q ”) satisfying the usual principles of propositional logic. We say p is *true* if \top entails p ; the context extension Γ, p hypothesizes that p is true.

Remark 2. Our proposition symbols are much more restricted than, and should not be confused with, other notions of proposition in type theory such as h-propositions (Univalent Foundations Program 2013, §3.3) or strict propositions (Gilbert et al. 2019). In particular, unlike types, our proposition symbols have no associated proof terms.

The type $\{p\} A$ is the dependent product “ $\{p\} A$ ” that is, $\{p\} A$ is well-formed when A is a type under the hypothesis that p is true, and $f : \{p\} A$ when, given that p is true, we may conclude $f : A$. The *extension type* $\{A|p \hookrightarrow a_p\}$ is well-formed when A is a type and $a_p : \{p\} A$; its elements $a : \{A|p \hookrightarrow a_p\}$ are terms $a : A$ satisfying the side condition that when p is true, we have $a = a_p : A$. We provide inference rules for the core calculus, including these connectives, in Section 4.1.

3.2 Elaborating controlled unfolding to our core calculus

Our surface language extends a generic surface language for dependent type theory with a new expression former **unfold** and several new declaration forms: ϑ **unfolds** $\kappa_1; \dots; \kappa_n$ for controlled unfolding, **abbreviation** ϑ for transparent definitions, and **abstract** ϑ for opaque definitions. Elaboration transforms these surface-language declarations into core-language signatures, that is, sequences of declarations over our core calculus of MLTT with proposition symbols.

Our signatures include the following declaration forms:

- **prop** $p \leq q$ introduces a fresh proposition symbol p such that p entails $q \in \mathbb{P}$;
- **prop** $p = q$ defines the proposition symbol p to be an abbreviation for $q \in \mathbb{P}$;
- **const** $\vartheta : A$ introduces a constant ϑ of type A .

We now revisit our examples from Section 2, illustrating how they are elaborated into our core calculus:

Plain definitions

Recall our unadorned definition of $(+)$ from Section 2:

$$(+) : \mathbb{N} \rightarrow \mathbb{N} \rightarrow \mathbb{N}$$

$$\text{ze } + n = n$$

$$\text{su } m + n = \text{su } (m + n)$$

We elaborate $(+)$ into a sequence of declarations: first, we introduce a new proposition symbol Υ_+ corresponding to the proposition that “ $(+)$ unfolds.” Next, we introduce a new definition $\delta_+ : \mathbb{N} \rightarrow \mathbb{N} \rightarrow \mathbb{N}$ satisfying the defining clauses of $(+)$ above, under the (trivial) assumption of \top ; finally, we introduce a new constant $(+)$ involving the extension type of δ_+ along Υ_+ :

$$\text{prop } \Upsilon_+ \leq \top$$

$$\delta_+ : \{\top\} (m n : \mathbb{N}) \rightarrow \mathbb{N}$$

$$\delta_+ \text{ ze } n = n$$

$$\delta_+ (\text{su } m) n = \text{su } (\delta_+ m n)$$

$$\text{const } (+) : \{\mathbb{N} \rightarrow \mathbb{N} \rightarrow \mathbb{N} | \Upsilon_+ \hookrightarrow \delta_+\}$$

Remark 3. In a serious implementation, it would be simple to induce δ_+ to be pretty-printed as $(+)$ in user-facing displays such as goals and error messages.

Top-level unfolding

To understand why we have elaborated $(+)$ in this way, let us examine how to elaborate top-level unfolding declarations (Section 2.1):

(\oplus) unfolds $(+)$

$(\oplus) : \text{vec } m \ A \rightarrow \text{vec } n \ A \rightarrow \text{vec } (m + n) \ A$

$[] \oplus v = v$

$(a :: u) \oplus v = a :: (u \oplus v)$

To elaborate **(\oplus) unfolds $(+)$** , we define the proposition symbol Υ_{\oplus} to entail Υ_+ , capturing the idea that unfolding (\oplus) always causes $(+)$ to unfold; in order to cause $(+)$ to unfold in the body of (\oplus) , we assume Υ_+ in the definition of δ_{\oplus} . In full, we elaborate the definition of (\oplus) as follows:

prop $\Upsilon_{\oplus} \leq \Upsilon_+$

$\delta_{\oplus} : \{\Upsilon_+\} (u : \text{vec } m \ A) (v : \text{vec } n \ A) \rightarrow \text{vec } (m + n) \ A$

$\delta_{\oplus} [] v = v$

$\delta_{\oplus} (a :: u) v = a :: (\delta_{\oplus} u \ v)$

const $(\oplus) : \{\text{vec } m \ A \rightarrow \text{vec } n \ A \rightarrow \text{vec } (m + n) \ A \mid \Upsilon_{\oplus} \hookrightarrow \delta_{\oplus}\}$

Observe that the definition of δ_{\oplus} is well-typed because Υ_+ is true in its scope: thus, the extension type of $(+)$ causes $\text{ze } + \ n$ to be definitionally equal to $\delta_+ \ \text{ze } n$, which in turn is defined to be n . The constraint $\Upsilon_{\oplus} \hookrightarrow \delta_{\oplus}$ is well-typed because Υ_{\oplus} entails Υ_+ .

If a definition ϑ unfolds multiple definitions $\kappa_1; \dots; \kappa_n$, we define Υ_{ϑ} to entail (and define δ_{ϑ} to assume) the conjunction $\Upsilon_{\kappa_1} \wedge \dots \wedge \Upsilon_{\kappa_n}$; if a definition ϑ unfolds no definitions, then Υ_{ϑ} entails (and δ_{ϑ} assumes) \top , as in our $(+)$ example.

Abbreviations

To elaborate the combination of the declarations **abbreviation** ϑ and ϑ **unfolds** $\kappa_1; \dots; \kappa_n$, we define Υ_{ϑ} to equal the conjunction $\Upsilon_{\kappa_1} \wedge \dots \wedge \Upsilon_{\kappa_n}$. For example, consider the following code from Section 2.3:

abbreviation $\text{map-}\oplus$

$\text{map-}\oplus$ **unfolds** $\text{map}; (\oplus)$

$\text{map-}\oplus : (f : A \rightarrow B) (u : \text{vec } m \ A) (v : \text{vec } n \ A) \rightarrow \text{map } f \ (u \oplus v) \equiv \text{map } f \ u \oplus \text{map } f \ v$

$\text{map-}\oplus \ f \ [] \ v = \text{refl}$

$\text{map-}\oplus \ f \ (a :: u) \ v = \text{cong } ((f \ a) :: -) \ (\text{map-}\oplus \ f \ u \ v)$

Let us write \mathfrak{C} for the following type:

$(f : A \rightarrow B) (u : \text{vec } m \ A) (v : \text{vec } n \ A) \rightarrow \text{map } f \ (u \oplus v) \equiv \text{map } f \ u \oplus \text{map } f \ v$

The above example is then elaborated as follows:

prop $\Upsilon_{\text{map-}\oplus} = \Upsilon_{\text{map}} \wedge \Upsilon_{\oplus}$

$\delta_{\text{map-}\oplus} : \{\Upsilon_{\text{map}} \wedge \Upsilon_{\oplus}\} \mathfrak{C}$

$\delta_{\text{map-}\oplus} f [] \nu = \text{refl}$

$\delta_{\text{map-}\oplus} f (a :: u) \nu = \text{cong } ((f \ a) :: -) (\delta_{\text{map-}\oplus} f \ u \ \nu)$

const $\text{map-}\oplus : \{\mathfrak{C} \mid \Upsilon_{\text{map-}\oplus} \hookrightarrow \delta_{\text{map-}\oplus}\}$

Expression-level unfolding

The elaboration of the expression-level unfolding construct **unfold** κ **in** M to our core calculus factors through the elaboration of expression-level unfolding to top-level unfolding as described in Section 2.5; we return to this in Section 4.3.

4. The Elaboration Algorithm

We now formally specify our mechanism for controlled unfolding by more precisely defining the elaboration algorithm sketched in the previous section, starting with a precise definition of the target of elaboration, our core calculus $\mathbf{TT}_{\mathbb{P}}$.

4.1 The core calculus $\mathbf{TT}_{\mathbb{P}}$

Our core calculus $\mathbf{TT}_{\mathbb{P}}$ is intensional MLTT (Martin-Löf 1975) with dependent sums and products, a Tarski universe, etc., extended with (1) a collection of proof-irrelevant proposition symbols, (2) dependent products over propositions, and (3) extension types for those propositions (Riehl and Shulman 2017).

Remark 4. We treat the features of MLTT and of our surface language somewhat generically; our elaboration algorithm can be applied on top of an existing bidirectional elaboration algorithm for type theory, for example, those described by Dagand (2013) and Gratzer et al. (2019), which may separately account for features such as implicit arguments or dependent pattern matching.

In fact, $\mathbf{TT}_{\mathbb{P}}$ is actually a family of type theories parameterized by a bounded meet semilattice $(\mathbb{P}, \top, \wedge)$ whose underlying set \mathbb{P} is the set of proposition symbols of $\mathbf{TT}_{\mathbb{P}}$; the semilattice structure on \mathbb{P} axiomatizes the conjunctive fragment of propositional logic with \wedge as conjunction, \top as the true proposition, and \leq as entailment (where $p \leq q$ is defined as $p \wedge q = p$), subject to the usual logical principles such as $p \wedge q \leq p$ and $p \wedge q \leq q$ and $p \leq \top$.

Remark 5. The judgments of $\mathbf{TT}_{\mathbb{P}}$ are functorial in the choice of \mathbb{P} , in the sense that given any homomorphism $f : \mathbb{P} \rightarrow \mathbb{P}'$ of bounded meet semilattices and any type or term in $\mathbf{TT}_{\mathbb{P}}$ over \mathbb{P} , we have an induced type/term in $\mathbf{TT}_{\mathbb{P}'}$ over \mathbb{P}' . In particular, we will use the fact that judgments of $\mathbf{TT}_{\mathbb{P}}$ are stable under adjoining new proposition symbols to \mathbb{P} .

The language $\mathbf{TT}_{\mathbb{P}}$ augments ordinary MLTT with a new judgment $\Gamma \vdash p \text{ true}$ (for $p \in \mathbb{P}$) and the corresponding context extension Γ, p (for $p \in \mathbb{P}$). The judgment $\Gamma \vdash p \text{ true}$ states that the

proposition p is true in context Γ , that is, the conjunction of the propositional hypotheses in Γ entails p while Γ, p extends Γ with the hypothesis that p is true.

$$\frac{\Gamma \text{ ctx} \quad p \in \mathbb{P}}{\Gamma, p \text{ ctx}} \quad \frac{p \in \mathbb{P}}{\Gamma, p \vdash p \text{ true}} \quad \frac{\Gamma, p \vdash \mathcal{J} \quad \Gamma \vdash p \text{ true}}{\Gamma \vdash \mathcal{J}} \\ \frac{}{\Gamma \vdash \top \text{ true}} \quad \frac{\Gamma \vdash p \text{ true} \quad \Gamma \vdash q \text{ true}}{\Gamma \vdash p \wedge q \text{ true}} \quad \frac{\Gamma \vdash p \text{ true} \quad p \leq q}{\Gamma \vdash q \text{ true}}$$

The dependent product $\{p\} A$ is defined as an ordinary dependent product:

$$\frac{\Gamma, p \vdash A \text{ type}}{\Gamma \vdash \{p\} A \text{ type}} \quad \frac{\Gamma, p \vdash M : A}{\Gamma \vdash \langle p \rangle M : \{p\} A} \quad \frac{\Gamma \vdash M : \{p\} A \quad \Gamma \vdash p \text{ true}}{\Gamma \vdash M @ p : A} \\ \frac{\Gamma, p \vdash M : A \quad \Gamma \vdash p \text{ true}}{\Gamma \vdash (\langle p \rangle M) @ p = M : A} \quad \frac{\Gamma \vdash M : pA}{\Gamma \vdash M = \langle p \rangle (M @ p) : \{p\} A}$$

The remaining feature of $\mathbf{TT}_{\mathbb{P}}$ is the extension type $\{A|p \hookrightarrow a_p\}$. Given a proposition $p \in \mathbb{P}$ and an element a_p of A under the hypothesis p , the elements of $\{A|p \hookrightarrow a_p\}$ correspond to elements of A that equal a_p when p holds.

$$\Gamma \vdash a : A \\ \Gamma, p \vdash a_p : A \\ \frac{\Gamma \vdash A \text{ type} \quad \Gamma, p \vdash a_p : A}{\Gamma \vdash \{A|p \hookrightarrow a_p\} \text{ type}} \quad \frac{\Gamma, p \vdash a = a_p : A}{\Gamma \vdash \text{in}_p a : \{A|p \hookrightarrow a_p\}} \quad \frac{\Gamma \vdash a : \{A|p \hookrightarrow a_p\}}{\Gamma \vdash \text{out}_p a : A} \\ \frac{\Gamma \vdash a : A}{\Gamma \vdash \text{out}_p(\text{in}_p a) = a : A} \quad \frac{\Gamma \vdash a : \{A|p \hookrightarrow a_p\}}{\Gamma \vdash \text{in}_p(\text{out}_p a) = a : \{A|p \hookrightarrow a_p\}} \\ \frac{\Gamma \vdash p \text{ true} \quad \Gamma \vdash a : \{A|p \hookrightarrow a_p\}}{\Gamma \vdash \text{out}_p a = a_p : A}$$

4.2 Signatures over $\mathbf{TT}_{\mathbb{P}}$

Our elaboration procedure takes as input a sequence of surface-language definitions and outputs a well-formed signature, a list of declarations over $\mathbf{TT}_{\mathbb{P}}$.

$$\begin{aligned} (\text{sigs}) \quad \Sigma &::= \epsilon \mid \Sigma, D \\ (\text{decls}) \quad D &::= \mathbf{const} \, x : A \mid \mathbf{prop} \, p \leq q \mid \mathbf{prop} \, p = q \end{aligned}$$

A signature is well-formed precisely when each declaration in Σ is well-formed relative to the earlier declarations in Σ . Our well-formedness judgment $\vdash \Sigma \text{ sig} \longrightarrow \mathbb{P}, \Gamma$ computes from Σ the $\mathbf{TT}_{\mathbb{P}}$ context Γ and proposition semilattice \mathbb{P} specified by Σ 's **const** and **prop** declarations, respectively.

The rules for signature well-formedness are standard except for the **prop** $p \leq q$ and **prop** $p = q$ declarations, which extend \mathbb{P} with a new element p satisfying $p \leq q$ or $p = q$, respectively. Recalling that our core calculus $\mathbf{TT}_{\mathbb{P}}$ is really a family of type theories parameterized by a semilattice \mathbb{P} , these declarations shift us between type theories, for example, from $\mathbf{TT}_{\mathbb{P}}$ to $\mathbf{TT}_{\mathbb{Q}}$, where $\mathbb{Q} = \mathbb{P}[p \leq q]$ is the minimal semilattice containing \mathbb{P} and an element p satisfying $p \leq q$. This shifting between theories is justified by Remark 5.

$$\begin{array}{c}
\frac{}{\vdash \epsilon \text{ sig} \longrightarrow \{\top\}, \cdot} \quad \frac{\vdash \Sigma \text{ sig} \longrightarrow \mathbb{P}, \Gamma \quad \Gamma \vdash_{\mathbf{TT}_{\mathbb{P}}} A \text{ type}}{\vdash (\Sigma, \mathbf{const} x : A) \text{ sig} \longrightarrow \mathbb{P}, (\Gamma, x : A)} \\
\frac{\vdash \Sigma \text{ sig} \longrightarrow \mathbb{P}, \Gamma \quad q \in \mathbb{P}}{\vdash (\Sigma, \mathbf{prop} p \leq q) \text{ sig} \longrightarrow \mathbb{P}[p \leq q], \Gamma} \quad \vdash (\Sigma, \mathbf{prop} p = q) \text{ sig} \longrightarrow \mathbb{P}[p = q], \Gamma
\end{array}$$

4.3 Bidirectional elaboration

We adopt a bidirectional elaboration algorithm which mirrors bidirectional type-checking algorithms (Coquand 1996; Pierce and Turner 2000). The top-level elaboration judgment $\Sigma \vdash \vec{S} \rightsquigarrow \Sigma'$ takes as input the current well-formed signature Σ and a list of surface-level definitions \vec{S} and outputs a new well-formed signature Σ' .

We define $\Sigma \vdash \vec{S} \rightsquigarrow \Sigma'$ in terms of three auxiliary judgments for elaborating surface-language types and terms; in the bidirectional style, we divide term elaboration into a checking judgment $\Sigma; \Gamma \vdash e \Leftarrow A \rightsquigarrow \Sigma', M$ taking a core type as input and a synthesis judgment $\Sigma; \Gamma \vdash e \Rightarrow A \rightsquigarrow \Sigma', M$ producing a core type as output. All three judgments take as input a signature Σ and a context (telescope) over Σ and output a new signature along with a core type or term.

We represent a surface-level definition S as a tuple:

$$(\mathbf{def} \vartheta : A, \text{abbrv?}, \text{abstr?}, [\kappa_1, \dots, \kappa_n], e)$$

In this expression, ϑ is the name of the definiendum, A is the surface-level type of the definition, abbrv? and abstr? are flags governing whether ϑ is an **abbreviation** (resp., is **abstract**), $[\kappa_1, \dots, \kappa_n]$ are the names of the definitions that ϑ unfolds, and e is the surface-level definiens.

The elaboration judgment elaborates each surface definition in sequence:

$$\begin{array}{l}
\Sigma \vdash \vec{S} \rightsquigarrow \Sigma_1 \\
\Sigma_1; \cdot \vdash A \Leftarrow \text{type} \rightsquigarrow \Sigma_2, A \\
\Sigma_2; \bigwedge_{i \leq n} \Upsilon_{\kappa_i} \vdash e \Leftarrow A \rightsquigarrow \Sigma_3, M \\
\mathbf{let} p := \mathbf{if} \text{abstr?} \mathbf{then} \text{gensym} () \mathbf{else} \Upsilon_{\vartheta} \\
\mathbf{let} \Box := \mathbf{if} \text{abbrv?} \mathbf{then} (=) \mathbf{else} (\leq) \\
\mathbf{let} \Sigma_4 := \Sigma_3, \mathbf{prop} p \Box \bigwedge_{i \leq n} \Upsilon_{\kappa_i}, \mathbf{const} \vartheta : \{A \mid p \hookrightarrow M\} \\
\hline
\Sigma \vdash \vec{S}, (\mathbf{def} \vartheta : A, \text{abbrv?}, \text{abstr?}, [\kappa_1, \dots, \kappa_n], e) \rightsquigarrow \Sigma_4
\end{array}$$

Remark 6. When a definition is marked **abstract**, the name of the unfolding proposition is generated fresh so that it cannot be accessed by any future **unfold** declaration. Conversely, when a definition is marked as an **abbreviation**, its unfolding proposition is defined to be equivalent to the conjunction of its dependencies rather than merely entailing its dependencies.

The rules for term and type elaboration are largely standard: for instance, we elaborate a surface-dependent product to a core-dependent product by recursively elaborating the first and second components. We single out two cases below: the boundary between checking and synthesis, and the expression-level **unfold**.

$$\begin{array}{c}
\Sigma; \Gamma \vdash e \Rightarrow A \rightsquigarrow \Sigma_1; M \quad \Sigma; \Gamma, \Upsilon_{\vartheta} \vdash e \Leftarrow A \rightsquigarrow \Sigma_1; M \quad \mathbf{let} \chi := \text{gensym} () \\
\hline
\Sigma_1; \Gamma \vdash \text{conv} A B \quad \mathbf{let} \Sigma_2 := \Sigma_1, \mathbf{const} \chi : \prod_{\Gamma} \{A \mid \Upsilon_{\vartheta} \hookrightarrow M\} \\
\hline
\Sigma; \Gamma \vdash e \Leftarrow B \rightsquigarrow \Sigma_1; M \quad \Sigma; \Gamma \vdash \mathbf{unfold} \vartheta \mathbf{in} e \Leftarrow A \rightsquigarrow \Sigma_2; \text{out}_{\Upsilon_{\vartheta}} \chi [\Gamma]
\end{array}$$

The first rule states that a term synthesizing a type A can be checked against a type B provided that A and B are definitionally equal; in order to implement this rule algorithmically, we need definitional equality to be decidable. Additionally, our (omitted) type-directed elaboration rules

are only well-defined if type constructors are injective up to definitional equality, for example, $A \rightarrow B = C \rightarrow D$ if and only if $A = C$ and $B = D$.

Elaborating expression-level unfolding requires the ability to hoist a type to the top level by iterating dependent products over its context, an operation notated \prod_{Γ} above. Because Γ can hypothesize (the truth of) propositions, this operation relies crucially on the presence of dependent products $\{p\} A$.

5. Case Study: An Implementation in **cooltt**

We have implemented our approach to controlled unfolding in the experimental **cooltt** proof assistant (RedPRL Development Team, 2020); **cooltt** is an implementation of cartesian cubical type theory (Angiuli *et al.* 2021), a computational version of homotopy type theory whose syntactic metatheory is particularly well understood (Huber 2019; Sterling 2021; Sterling and Angiuli 2021). The existing support for partial elements and extension types made **cooltt** particularly hospitable for experimentation with elaborating controlled unfolding to extension types. The following example illustrates the use of controlled unfolding in **cooltt**, where $\text{path } A \ x \ y$ is the cubical notion of propositional equality ($x \equiv y$):

```

def + :  $\mathbb{N} \rightarrow \mathbb{N} \rightarrow \mathbb{N} :=
  \text{elim}
  | \text{zero} \Rightarrow n \Rightarrow n
  | \text{suc } \_ \Rightarrow ih \Rightarrow n \Rightarrow \text{suc } \{ih\} n

unfold +
def +0L ( $x : \mathbb{N}$ ) :  $\text{path } \mathbb{N} \ \{+ \ 0 \ x\} \ x :=
  i \Rightarrow x

def +0R : ( $x : \mathbb{N}$ )  $\rightarrow$   $\text{path } \mathbb{N} \ \{+ \ x \ 0\} \ x :=
  \text{elim}
  | \text{zero} \Rightarrow +0L \ 0
  | \text{suc } \{x \Rightarrow ih\} \Rightarrow
    \text{equation } \mathbb{N}
    | + \ 0 \ \{\text{suc } y\} = [+0L \ \{\text{suc } y\}]
    | \text{suc } \{+ \ x \ 0\} = [i \Rightarrow \text{suc } \{ih\} i]
    | \text{suc } x$$$ 
```

This example follows a common pattern: we prove basic computational laws (+0L) by unfolding a definition, and then in subsequent results (+0R) use these lemmas abstractly rather than unfolding. Doing so controls the size and readability of proof goals and explicitly demarcates which parts of the library depend on the definitional behavior of a given function.

We have also implemented the derived forms for expression-level unfolding:

```

def two :  $\mathbb{N} := + \ 1 \ 1

def thm :  $\text{path } \mathbb{N} \ \text{two} \ 2 := \text{unfold } \text{two} + \text{in } i \Rightarrow 2

def thm-is-refl :  $\text{path-p } \{i \Rightarrow \text{path } \mathbb{N} \ \text{two} \ \{\text{thm } i\}\} \ \{i \Rightarrow \text{two}\} \ \text{thm} :=
  i \ j \Rightarrow \text{unfold } \text{two} + \text{in } 2$$$ 
```

def thm-is-refl' : path {path \mathbb{N} two 2} { $i \Rightarrow$ **unfold** two + **in** two} thm :=
 $i\ j \Rightarrow$ **unfold** two + **in** 2

The third and fourth declarations above illustrate two strategies in **cooltt** for dealing with a dependent type whose well-formedness depends on an unfolding; in **thm-is-refl**, we use a dependent path type but only unfold in the definiens, whereas in **thm-is-refl'** we use a non-dependent path type but must unfold in both the definiens and in its type.

Our **cooltt** implementation deviates in a few respects from the presentation in this paper: in particular, the propositions Υ_k are represented by abstract elements $i_k : \mathbb{I}$ of the interval via the embedding $\mathbb{I} \hookrightarrow \mathbb{F}$ sending i to $(i =_{\mathbb{I}} 1)$.

cooltt utilizes a standalone library to compute entailment of cofibrations called Kado (Hou (Favonia) 2022), created by Kuen-Bang Hou (Favonia). To support our experiment, Favonia modified Kado to support inequalities of dimension variables $i \leq_{\mathbb{I}} j$ in addition to the cofibrations needed for **cooltt**'s core theory. As a result, the modifications to **cooltt** were quite modest. After the changes to Kado – which could in principle be reused in any proof assistant for the same purpose – the entire change required only a net increase of 996 lines of OCaml code.

6. The Metatheory of $\mathbf{TT}_{\mathbb{P}}$

In Section 4, we described an algorithm elaborating a surface language with controlled unfolding to $\mathbf{TT}_{\mathbb{P}}$. In order to actually execute our algorithm, it is necessary to decide the definitional equality of types in $\mathbf{TT}_{\mathbb{P}}$; as is often the case in type theory, type dependency ensures that deciding equality for types also requires us to decide the equality of terms. In order to implement our elaboration algorithm, we therefore prove a normalization result for $\mathbf{TT}_{\mathbb{P}}$.

At its heart, a normalization algorithm is a computable bijection between equivalence classes of terms up to definitional equality and a collection of normal forms. By ensuring that the equality of normal forms is evidently decidable, this yields an effective decision procedure for definitional equality. In our case, we attack normalization through a synthetic and semantic approach to normalization by evaluation called STC (Sterling and Angiuli 2021; Sterling, 2021, 2025; Sterling and Harper 2021).

Neutral forms for $\mathbf{TT}_{\mathbb{P}}$. The semantic analysis of normalization by evaluation rests on the observation of Fiore (2002) that normal forms, though not stable under arbitrary substitutions, are nonetheless stable under renamings – substitutions that replace variables with variables (not necessarily injective). Therefore, decisive aspects of the normalization algorithm can be expressed internally to a topos of variable sets (presheaves) over the category of contexts and renamings; in order to instrument the semantic normalization algorithm with its proof of correctness, one passes to a larger topos obtained from the former by gluing. STC then instantiates the standard topos model of MLTT to substantially simplify various details that would otherwise be exceedingly tedious by means of a form of higher-order abstract syntax.

This appealingly simple story for normalization is substantially complicated by the boundary law for extension types:

$$\frac{\Gamma \vdash p \text{ true} \quad \Gamma, p \vdash a_p : A \quad \Gamma \vdash a : \{A \mid p \hookrightarrow a_p\}}{\Gamma \vdash \text{out}_p a = a_p : A}$$

When defining normal forms for $\mathbf{TT}_{\mathbb{P}}$, we might naively add a neutral form out_p to represent out_p . In order to ensure that normal and neutral forms correspond bijectively with equivalence classes of terms, however, we should only allow out_p to be applied in a context where p is not true; if p were true, $\text{out}_p a$ is already represented by the normal form for a_p .

A similar problem arises in the context of cubical type theory (Angiuli et al. 2021; Cohen et al. 2017) where some equalities apply precisely when two dimensions coincide. The same problem

arises: either renamings must exclude substitutions that identify two dimension terms, or neutral forms will not be stable under renamings. In their proof of normalization for cubical type theory, Sterling and Angiuli (2021) refined neutral forms to account for this tension by introducing stabilized neutrals. Rather than cutting down on renamings, they expand the class of neutrals by allowing “bad” neutrals akin to $\text{out}_p e$ in a context where p is true. They then associate each neutral form with a frontier of instability: a proposition that becomes true when the neutral is no longer “stuck.” Crucially, although well-behaved neutrals may not be stable under renamings, the frontier of instability is stable and can therefore be incorporated into the internal language.

We adapt Sterling and Angiuli’s stabilized neutrals to the simplified setting of $\mathbf{TT}_{\mathbb{P}}$ and establish its normalization theorem. In so doing, we refine the approach of *op. cit.* to obtain a fully constructive³ normalization proof. We also carefully spell out the details of the universe in the normalization model, correcting an oversight in an earlier revision of Sterling’s dissertation (Sterling 2021).

6.1 Type theories as categories with representable maps

While any number of logical frameworks are available (generalized algebraic theories (Cartmell 1978), essentially algebraic theories (Freyd 1972), locally cartesian closed categories (Gratzer and Sterling 2020), etc.), Uemura’s categories with representable maps (Uemura, 2021, 2023) are particularly attractive because they express exactly the binding and dependency structure needed for type theory: a second-order version of generalized algebraic theories.

Definition 7. A category with representable maps (CwR) \mathcal{C} is a finitely complete category equipped with a pullback-stable class of representable maps $\mathcal{R} \subseteq \text{Arr}(\mathcal{C})$ such that pullback along $f \in \mathcal{R}$ has a right adjoint (dependent product along f).

Definition 8. A morphism of CwRs is a functor between the underlying categories that preserves finite limits, representability of maps, and dependent products along representable maps.

Definition 9. CwRs, morphisms between them, and natural isomorphisms assemble into a $(2, 1)$ -category \mathbf{CwR} .

Uemura’s logical framework axiomatizes the category of judgments of $\mathbf{TT}_{\mathbb{P}}$ as a particular category with representable maps \mathbb{T} . The finite limit structure of \mathbb{T} encodes substitution as well as equality judgments, while the class of representable maps carves out those judgments that may be hypothesized. Uemura (2023) develops a syntactic method for presenting a CwR as a signature within a variant of extensional type theory, which he has rephrased in terms of second-order generalized algebraic theories in his doctoral dissertation (Uemura 2021). Although we will use the type-theoretic presentation for convenience, the difference between these two accounts is only superficial.

Each judgment of $\mathbf{TT}_{\mathbb{P}}$ is rendered as a (dependent) sort, while operators are modeled by elements of the given sorts. In order to record whether a given judgment may be hypothesized, the sorts of the type theory are stratified by meta-sorts $\star \subseteq \square$ where $A : \star$ signifies that A is a representable sort (i.e., a context-former) and can be hypothesized, whereas $B : \square$ cannot parameterize a framework-level dependent product.

Proposition 10. Let \mathbb{T} be the free category with representable maps generated by a given logical framework signature; then the groupoid of CwR functors $\text{hom } \mathbf{CwR} \mathbb{T} \mathcal{E}$ is equivalent to the groupoid of interpretations of the signature within \mathcal{E} .

We will often refer to a category with representable maps \mathbb{T} as a type theory; indeed, as the category of judgments of a given type theory, \mathbb{T} is a suitable invariant replacement for it.

Proposition 10 describes the universal property of a type theory generated by a given signature in a logical framework. Type theories qua CwRs thus give rise to a form of functorial semantics in which algebras (interpretations) arrange into a groupoid of CwR functors $\mathbf{hom} \mathbf{CwR} \mathbb{T} \mathcal{C}$.

This is an appropriate setting for studying the syntax of type theory, but it is somewhat inappropriate for studying the semantics of type theory – in which one expects models to correspond to structured CwFs (Dybjer 1996) or natural models (Awodey 2018), which themselves arrange into a (2,1)-category. The second notion of functorial semantics, developed by Uemura in his doctoral dissertation (Uemura 2021), is a generalization of the theory of CwFs and pseudo-morphisms between them (Clairambault and Dybjer 2014; Newstead 2018).

Note that we may always regard a presheaf category $\mathbf{Pr} \mathcal{C}$ as a CwR with the representable maps being representable natural transformations, that is, families of presheaves whose fibers at representables are representable (Awodey 2018).

Definition 11. A model of a type theory \mathbb{T} is a category \mathbf{M}_\diamond together with a CwR functor $\mathbf{M} : \mathbb{T} \rightarrow \mathbf{Pr} \mathbf{M}_\diamond$.

Models are arranged into a (2,1)-category $\mathbf{Mod} \mathbb{T}$ (see Appendix A). Essentially, a morphism of models $\mathbf{M} \rightarrow \mathbf{N}$ is given by a functor $\alpha_\diamond : \mathbf{M}_\diamond \rightarrow \mathbf{N}_\diamond$ together with a natural transformation $\mathbf{M} \rightarrow \alpha^* \mathbf{N} \in \mathbf{hom}_{\mathbf{CwR}}(\mathbb{T}, \mathbf{Pr} \mathbf{M}_\diamond)$ that preserves context extensions up to isomorphism; an isomorphism between morphisms of models is a natural isomorphism between the underlying functors satisfying an additional property.

For each CwR \mathbb{T} , Uemura has shown the following theorem:

Proposition 12. The (2,1)-category of models $\mathbf{Mod} \mathbb{T}$ has a bi-initial object \mathbf{I} , whose category of contexts \mathbf{I}_\diamond is the smallest full subcategory of \mathbb{T} closed under the terminal object and pullbacks along representable maps.

Remark 13. If one takes \mathbb{T} to be for example, MLTT, the bi-initial model \mathbf{I} can be realized by the familiar initial CwF built from the category of contexts.

6.2 Encoding $\mathbf{TT}_{\mathbb{P}}$ in the logical framework

We begin by defining the signature for a category with representable maps \mathbb{T}_0 containing exactly the bare judgmental structure of $\mathbf{TT}_{\mathbb{P}}$, namely the propositions and the judgments for types and terms. In our signature, we make liberal use of the Agda-style notation for implicit arguments. As always, p ranges over \mathbb{P} :

$\langle p \rangle : \star$

$\mathbf{tp} \square$

$\mathbf{tm} : \mathbf{tp} \Longrightarrow \star$

$_ : \{u, v : \langle p \rangle\} \Longrightarrow u = v$

$_ : \{_ : \langle \bigwedge_{i < n} p_i \rangle\} \Longrightarrow \langle p_k \rangle$

$_ : \{_ : \langle p_i \rangle, \dots\} \Longrightarrow \langle \bigwedge_{i < n} p_i \rangle$

Note that already this signature encodes the necessary theory of propositions for $\mathbf{TT}_{\mathbb{P}}$. For instance, if $p \leq q$ in \mathbb{P} , then a combination of the final two implications in the signature implies $\langle p \rangle \rightarrow \langle q \rangle$. We next extend the above to include the type formers of $\mathbf{TT}_{\mathbb{P}}$, writing \mathbb{T} for the CwR generated by the full signature.

$$\begin{aligned}
& \text{tp} : \square \\
& \text{tm} : \text{tp} \Longrightarrow \star \\
& \langle p \rangle : \star \\
& _ : \{u, v : \langle p \rangle\} \Longrightarrow u = v \\
& _ : \{ _ : \langle \bigwedge_{i < n} p_i \rangle \} \Longrightarrow \langle p_k \rangle \\
& _ : \{ _ : \langle p_i \rangle, \dots \} \Longrightarrow \langle \bigwedge_{i < n} p_i \rangle \\
& \text{ext}_p : (A : \text{tp}) (a : \{p\} \text{tm } A) \Longrightarrow \text{tp} \\
& \text{in}_p : (A : \text{tp}) (a : \{p\} \text{tm } A) (u : \text{tm } A) \{ _ : \{p\} u = a \} \Longrightarrow \text{tm} (\text{ext}_p A a) \\
& \text{out}_p : (A : \text{tp}) (a : \{p\} \text{tm } A) (u : \text{tm} (\text{ext}_p A a)) \Longrightarrow \text{tm } A \\
& _ : (A : \text{tp}) (a : \{p\} \text{tm } A) (u : \text{tm} (\text{ext}_p A a)) \{ _ : \langle p \rangle \} \Longrightarrow \text{out}_p A a u = a \\
& _ : (A : \text{tp}) (a : \{p\} \text{tm } A) (u : \text{tm } A) \{ _ : \{p\} u = a \} \Longrightarrow \text{out}_p A a (\text{in}_p A a u) = u \\
& _ : (A : \text{tp}) (a : \{p\} \text{tm } A) (u : \text{tm} (\text{ext}_p A a)) \Longrightarrow \text{in}_p A a (\text{out}_p A a u) = u \\
& \text{part}_p : (A : \{p\} \text{tp}) \Longrightarrow \text{tp} \\
& \text{lam}_p : (A : \text{tp}) (a : \{p\} \text{tm } A) \Longrightarrow \text{tm} (\text{part}_p A) \\
& \text{app}_p : (A : \text{tp}) (u : \text{tm} (\text{part}_p A)) \Longrightarrow \{p\} \text{tm } A \\
& _ : (A : \text{tp}) (a : \{p\} \text{tm } A) \Longrightarrow \text{app}_p A (\text{lam}_p A a) = a \\
& _ : (A : \text{tp}) (a : \text{tm} (\text{part}_p A)) \Longrightarrow \text{lam}_p A (\text{app}_p A a) = a
\end{aligned}$$

Figure 1. The non-standard aspects of the LF signature for $\mathbf{TT}_{\mathbb{P}}$.

Notation 14. Given $X : \{ _ : \langle p \rangle \} \Longrightarrow \square$, we will write $\{p\} X$ to further abbreviate the Agda-style implicit function space $\{ _ : \langle p \rangle \} \rightarrow X$. Note that $\{p\} X$ still associates with the right and so $\{p\} A \rightarrow B$ signifies $\{p\}(A \rightarrow B)$.

For instance, the following constants specify the rules of extension types given in Section 4.1:

$$\begin{aligned}
& \text{ext}_p : (A : \text{tp}) (a : \{p\} \text{tm } A) \Longrightarrow \text{tp} \\
& \text{in}_p : (A : \text{tp}) (a : \{p\} \text{tm } A) (u : \text{tm } A) \{ _ : \{p\} u = a \} \Longrightarrow \text{tm} (\text{ext}_p A a) \\
& \text{out}_p : (A : \text{tp}) (a : \{p\} \text{tm } A) (u : \text{tm} (\text{ext}_p A a)) \Longrightarrow \text{tm } A \\
& _ : (A : \text{tp}) (a : \{p\} \text{tm } A) (u : \text{tm} (\text{ext}_p A a)) \{ _ : \langle p \rangle \} \Longrightarrow \text{out}_p A a u = a \\
& _ : (A : \text{tp}) (a : \{p\} \text{tm } A) (u : \text{tm } A) \{ _ : \{p\} u = a \} \Longrightarrow \text{out}_p A a (\text{in}_p A a u) = u \\
& _ : (A : \text{tp}) (a : \{p\} \text{tm } A) (u : \text{tm} (\text{ext}_p A a)) \Longrightarrow \text{in}_p A a (\text{out}_p A a u) = u
\end{aligned}$$

The full list of non-standard constants is specified in Figure 1. Once the signature is complete, we obtain from Uemura's framework a category with representable maps \mathbb{T} together with a bi-initial model **I**.

6.3 The atomic figure shape and its universal property

For each context Γ and type $\Gamma, \Gamma \vdash A$ type, it is possible to axiomatize the normal forms of type A ; unfortunately, this assignment of sets of normal forms does not immediately extend to a presheaf on the category of contexts \mathbf{I}_\diamond , precisely because normal forms are not a priori closed under substitution. In fact, closing normal forms under substitution is the purpose of normalization, so we are not able to assume it beforehand.

Normal forms are, however, closed under substitutions of variables for variables (often called structural renamings), and in our case we shall be able to close them additionally under the “phase

transitions” $\Gamma, \langle p \rangle \rightarrow \Gamma, \langle q \rangle$ when $\Gamma, p \vdash q$ true is derivable. We shall refer to these substitutions as *atomic substitutions*, and we wish to organize them into a category.

It is possible to inductively define a category of “atomic contexts” whose objects are those of \mathbf{I}_\diamond and whose morphisms are atomic substitutions, but this construction obscures a beautiful and simple (2,1)-categorical universal property first exposed by Bocquet et al. (2021) that leads to a more modular proof. To explicate this universal property, first note that the theory \mathbb{T}_0 axiomatizes exactly the structure of variables and phase transitions, and that the initial model \mathbf{I} of \mathbb{T} is, by restriction along $\mathbb{T}_0 \rightarrow \mathbb{T}$, also a model of \mathbb{T}_0 .

Definition 15. An atomic substitution model over a fixed \mathbb{T} -model \mathbf{M} is given by a model \mathbf{A} of the bare judgmental theory \mathbb{T}_0 , together with a morphism of models $\alpha : \mathbf{A} \rightarrow \mathbf{M}$ in $\mathbf{Mod} \mathbb{T}_0$ such that $\alpha_{\text{tp}} : \mathbf{A}(\text{tp}) \rightarrow \alpha^*(\mathbf{M}(\text{tp})) \in \mathbf{Pr} \mathbf{A}_\diamond$ is an isomorphism.

Atomic substitution model over \mathbf{M} arrange themselves into a (2,1)-category, a full subcategory of $\mathbf{Mod} \mathbb{T}_0 \downarrow \mathbf{M}$. The following result is due to Bocquet et al. (2021).

Proposition 16. The bi-initial atomic substitution model $(\mathbf{A}, \alpha : \mathbf{A} \rightarrow \mathbf{I})$ over \mathbf{I} exists.

When $(\mathbf{A}, \alpha : \mathbf{A} \rightarrow \mathbf{I})$ is the bi-initial atomic substitution model over \mathbf{I} as in Proposition 16, we shall refer to an object $\Gamma \in \mathbf{A}$ as an atomic context and a morphism $\gamma : \Delta \rightarrow \Gamma$ in \mathbf{A} as an atomic substitution. We shall assume without loss of generality that $\mathbf{A}(\text{tp}) = \alpha^*\mathbf{I}(\text{tp})$ so that the component α_{tp} is the identity map.

6.4 Computability spaces by gluing along the atomic figure shape

We shall use the bi-initial atomic substitution model over \mathbf{I} as a *figure shape* in the sense of Sterling (2021, §4.3) to instantiate STC. Here, we transition into the 2-category of Grothendieck topoi, geometric morphisms, and geometric transformations, guided by a *phase distinction* between “object-space” and “meta-space” (Sterling 2021);⁴ object-space refers to the object language embodied in the model \mathbf{I} , whereas meta-space refers to the metalanguage embodied in the model \mathbf{A} . Later on, we will construct a glued topos in which we may speak of constructs that have extent in both object-space and meta-space. We follow Vickers (2007) and Anel and Joyal (2021) in emphasizing the distinction between a topos \mathbf{x} and the category of sheaves $\mathbf{Sh} \mathbf{x}$ presenting it:

Definition 17. We denote by \mathbf{I} and \mathbf{A} the object-space and meta-space topoi, respectively, with underlying categories of sheaves $\mathbf{Sh} \mathbf{I} = \mathbf{Pr} \mathbf{I}_\diamond$ and $\mathbf{Sh} \mathbf{A} = \mathbf{Pr} \mathbf{A}_\diamond$.

Definition 18. The functor $\alpha_\diamond : \mathbf{A}_\diamond \rightarrow \mathbf{I}_\diamond$ gives rise under precomposition to a continuous and cocontinuous functor $\mathbf{Pr} \mathbf{I}_\diamond \rightarrow \mathbf{Pr} \mathbf{A}_\diamond$ that shall serve as the inverse image part of an (essential) geometric morphism $\alpha : \mathbf{A} \rightarrow \mathbf{I}$ named the atomic figure shape.

That $\alpha : \mathbf{A} \rightarrow \mathbf{I}$ is essential means that its inverse image $\alpha^* : \mathbf{Sh} \mathbf{I} \rightarrow \mathbf{Sh} \mathbf{A}$ has a left adjoint $\alpha_! : \mathbf{Sh} \mathbf{A} \rightarrow \mathbf{Sh} \mathbf{I}$; from the point of view of presheaves, this is precisely the Yoneda extension of $\alpha_\diamond : \mathbf{A}_\diamond \rightarrow \mathbf{I}_\diamond$ as depicted below:

$$\begin{array}{ccc}
 \mathbf{A}_\diamond & \xrightarrow{\alpha_\diamond} & \mathbf{I}_\diamond \\
 \downarrow \gamma_{\mathbf{A}_\diamond} & & \downarrow \gamma_{\mathbf{I}_\diamond} \\
 \mathbf{Pr} \mathbf{A}_\diamond & \dashrightarrow_{\alpha_!} & \mathbf{Pr} \mathbf{I}_\diamond
 \end{array}$$

Definition 19. We denote by \mathbf{G} the closed mapping cylinder (Johnstone 1977) of the geometric morphism $\alpha : \mathbf{A} \rightarrow \mathbf{I}$; in other words, $\mathbf{Sh} \mathbf{G}$ is the comma category $\mathbf{Sh} \mathbf{A} \downarrow \alpha^*$. We will write $j : \mathbf{I} \hookrightarrow \mathbf{G}$ and $i : \mathbf{A} \hookrightarrow \mathbf{G}$ for the open and closed subtopos immersions.

Following Sterling (2025), we shall refer to a sheaf on \mathbf{G} as a *computability space*. A computability space $X \in \mathbf{Sh} \mathbf{G}$ is then identified with a family $\pi_X : i^*X \rightarrow \alpha^*j^*X$ in $\mathbf{Sh} \mathbf{A}$. Because the assignment π_X is natural in computability spaces X , it corresponds to a 2-cell $\pi : j \circ \alpha \rightarrow i$ in the 2-category of Grothendieck topoi. The universal property of \mathbf{G} is then expressed by the fact that $\pi : j \circ \alpha \rightarrow i$ is a co-comma cell in the 2-category of Grothendieck topoi:

$$\begin{array}{ccc} \mathbf{A} & \xrightarrow{\alpha} & \mathbf{I} \\ \parallel & \uparrow \pi & \downarrow j \\ \mathbf{A} & \xrightarrow{i} & \mathbf{G} \end{array}$$

Remark 20 (Relation to Kripke computability predicates). Unraveling Definition 19, a computability space is precisely a family X' of presheaves on \mathbf{A}_\diamond indexed in the restriction of a given presheaf X on \mathbf{I}_\diamond along $\alpha_\diamond : \mathbf{A}_\diamond \rightarrow \mathbf{I}_\diamond$. When the family X' is valued in subterminal presheaves and the base X is representable, we have precisely the classical notion of a Kripke computability predicate (Jung and Tiuryn 1993); a computability space in our sense is then a generalized, proof-relevant version of a Kripke computability predicate.

6.4.1 Reflection of object and meta-space

By definition, the inverse image functors j^*, i^* have fully faithful right adjoints $j_* : \mathbf{Sh} \mathbf{I} \hookrightarrow \mathbf{Sh} \mathbf{G}$ and $i_* : \mathbf{Sh} \mathbf{A} \hookrightarrow \mathbf{Sh} \mathbf{G}$, respectively. These are computed as follows:

$$\begin{aligned} j_*E &= (E, 1_{\alpha^*E} : \alpha^*E \rightarrow \alpha^*E) \\ i_*A &= (\mathbf{1Sh} \mathbf{I}, !_A : A \rightarrow \mathbf{Sh} \mathbf{A} \cong \alpha^*\mathbf{Sh} \mathbf{I}) \end{aligned}$$

Thus, the adjunctions $j^* \dashv j_*$ and $i^* \dashv i_*$ exhibit $\mathbf{Sh} \mathbf{I}$ and $\mathbf{Sh} \mathbf{A}$ as reflective subcategories of $\mathbf{Sh} \mathbf{G}$:

- (1) The essential image of the reflective embedding $j_* : \mathbf{Sh} \mathbf{I} \hookrightarrow \mathbf{Sh} \mathbf{G}$ is spanned by computability spaces X for which $\pi_X : i^*X \rightarrow \alpha^*j^*X$ is an isomorphism, that is, such that π_X is terminal in the slice $\mathbf{Sh} \mathbf{A} \downarrow \alpha^*j^*X$.
- (2) The essential image of the reflective embedding $i_* : \mathbf{Sh} \mathbf{A} \hookrightarrow \mathbf{Sh} \mathbf{G}$ is spanned by computability spaces X such that j^*X is terminal in $\mathbf{Sh} \mathbf{I}$.

Definition 21 (Vocabulary for reflective subcategories). When a computability space lies in the essential image of $j_* : \mathbf{Sh} \mathbf{I} \hookrightarrow \mathbf{Sh} \mathbf{G}$, we shall refer to it as lying in object-space. Likewise, when a computability space lies in the essential image of $i_* : \mathbf{Sh} \mathbf{A} \hookrightarrow \mathbf{Sh} \mathbf{G}$, we shall say that it lies in meta-space.

6.4.2 Coreflection of object- and meta-space

Both the open and closed immersions are essential morphisms of topoi, in the sense that we have additional (necessarily fully faithful) left adjoints $j_! \dashv j^* : \mathbf{Sh} \mathbf{G} \rightarrow \mathbf{Sh} \mathbf{I}$ and $i_! \dashv i^* : \mathbf{Sh} \mathbf{G} \rightarrow \mathbf{Sh} \mathbf{A}$

that are computed as follows:

$$\begin{aligned} j_!E &= (E, \downarrow_{\alpha^*E} : \mathbf{Sh} \mathbf{A} \rightarrow \alpha^*E) \\ i_!A &= (\alpha_!A, \eta_A : A \rightarrow \alpha^*\alpha_!A) \end{aligned}$$

Thus, $\mathbf{Sh} \mathbf{I}$ and $\mathbf{Sh} \mathbf{A}$ are not only reflective in $\mathbf{Sh} \mathbf{G}$ – they are also coreflective.

6.5 The language of STC

As \mathbf{I} and \mathbf{A} are both subtopoi of \mathbf{G} , their reflections (Section 6.4.1) can be expressed in the internal language of $\mathbf{Sh} \mathbf{G}$ by means of a pair of complementary lex idempotent monads (\circ, \bullet) . The internal language of $\mathbf{Sh} \mathbf{I}$ is presented by the \circ -modal or object-space types and $\mathbf{Sh} \mathbf{A}$ is presented by \bullet -modal or meta-space types. Because they form an open/closed partition, these modal subuniverses admit a particularly simple formulation:

Theorem 22. *There exists a proposition $\text{obj} : \Omega$ such that*

- (1) *a type X is \circ -modal/object-space iff $X \rightarrow (\text{obj} \rightarrow X)$ is an isomorphism;*
- (2) *a type X is \bullet -modal/meta-space iff $\text{obj} \times X \rightarrow \text{obj}$ is an isomorphism.*

Remark 23. In fact, the coreflection of $\mathbf{Sh} \mathbf{I}$ in $\mathbf{Sh} \mathbf{G}$ lifts smoothly into the internal language (though we shall not use this fact) by the idempotent comonadic modality $\Box \dashv \circ$ that sends X to the product $\Box X = \text{obj} \times X$. On the other hand, the coreflection of $\mathbf{Sh} \mathbf{A}$ in $\mathbf{Sh} \mathbf{G}$ cannot be expressed directly in the internal language.

Notation 24. *We will use extension types $\{A|\phi \hookrightarrow a\}$ in the internal language of $\mathbf{Sh} \mathbf{G}$ as realized by the subset comprehension of topos logic, treating their introduction and elimination rules silently. Here, ϕ will be an element of the subobject classifier, in contrast to the situation in our object language, where it ranged over fixed proposition symbols.*

Remark 25. We assume a subuniverse $\Omega_{\text{dec}} \subseteq \Omega$ of the subobject classifier that is closed under finite disjunctions and contains obj ; then Ω_{dec} will ultimately be a subuniverse spanned by pointwise/externally decidable propositions (Angiuli et al. 2021), but this fact will not play a role in the synthetic development.

Notation 26. *We will reuse Notation 14 and write $\{\text{obj}\} A$ rather than $\{_ : \text{obj}\} \rightarrow A$ when $A : \{_ : \text{obj}\} \rightarrow \mathcal{U}$.*

As a presheaf topos, $\mathbf{Sh} \mathbf{G}$ inherits a hierarchy of cumulative universes \mathcal{U}_i , each of which supports the *strict gluing* or (*mixed-phase*) *refinement* type (Gratzer et al. 2022): a version of the dependent sum of a family of meta-space types indexed in an object-space type A that additionally restricts within object-space to exactly A :

$$\frac{A : \{\text{obj}\} \mathcal{U}_i \quad B : (\{\text{obj}\} A) \rightarrow \mathcal{U}_i \quad \{\text{obj}\}(a : A) \rightarrow (B a \cong)}{(x : A) \ltimes B x : \{\mathcal{U}_i | \text{obj} \hookrightarrow A\} \quad \text{gl} : \{((x : \{\text{obj}\} A) \times B x) \cong (x : A) \ltimes B x | \text{obj} \hookrightarrow \pi_1\}}$$

Remark 27. In topos logic, it is a property for a function to have an inverse; thus, we have conveniently packaged the introduction and elimination rules for $(x : A) \ltimes B x$ into a single function gl that is assumed to be an isomorphism.

Notation 28. *We write $[\text{obj} \hookrightarrow a|b]$ for $\text{gl}(a, b)$ and $\text{ungl } x$ for $\pi_2(\text{gl}^{-1}x)$. When constructing particularly complex inhabitants of $(x : A) \ltimes B x$, we will avail ourselves of copattern matching notation*

and write the following instead of $c = [\text{obj} \hookrightarrow a|b]$:

$$\text{obj} \hookrightarrow c = a$$

$$\text{ungl } c = b$$

Both \circ and \bullet induce reflective subuniverses $\mathcal{U}_\circ^i, \mathcal{U}_\bullet^i \hookrightarrow \mathcal{U}_i$ spanned by modal types, and these universes are themselves modal. Following Sterling (2021), we use strict gluing to choose these universes with additional strict properties:

$$\mathcal{U}_\circ^i : \{\mathcal{U}_{i+1} | \text{obj} \hookrightarrow \mathcal{U}_i\} \quad \mathcal{U}_\bullet^i : \{\mathcal{U}_{i+1} | \text{obj} \hookrightarrow \}$$

Furthermore, the inclusion $\mathcal{U}_\circ^i \hookrightarrow \mathcal{U}_i$ restricts to the identity under obj . With the modal universes to hand, we may choose $\circ : \mathcal{U}_i \rightarrow \mathcal{U}_i$ and $\bullet : \mathcal{U}_i \rightarrow \mathcal{U}_i$ to factor through \mathcal{U}_\circ^i and \mathcal{U}_\bullet^i , respectively. Henceforth, we will suppress the inclusions $\mathcal{U}_\circ^i, \mathcal{U}_\bullet^i \hookrightarrow \mathcal{U}_i$ and write, for example, $\circ : \mathcal{U}_i \rightarrow \mathcal{U}_\circ^i$ for the reflections.

Remark 29. The strict gluing types, modal universes, and their modal reflections can be chosen to commute strictly with the liftings $\mathcal{U}_i \rightarrow \mathcal{U}_{i+1}$.

The interpretation of the $\mathbf{TT}_{\mathbb{P}}$ signature within $\mathbf{Sh} \mathbf{I}$ internalizes into $\mathbf{Sh} \mathbf{G}$ as a sequence of constants valued in the subuniverse \mathcal{U}_\circ^0 ; for instance, we have:

$$\text{tp} : \mathcal{U}_\circ^0$$

$$\text{tm} : \text{tp} \mathcal{U}_\circ^0$$

$$\langle p \rangle : \Omega_{\text{dec}} \text{ (for } p \in \mathbb{P} \text{)}$$

$$\text{extp} : (A : \text{tp}) \rightarrow (a : \{\langle p \rangle\} \text{tm} A) \rightarrow \text{tp}$$

$$\text{in}_p : (A : \text{tp}) (a : \{\langle p \rangle\} \text{tm} A) \rightarrow \{\text{tm} A | \langle p \rangle \hookrightarrow a\} \cong \text{tm}(\text{extp } A \ a)$$

Following Remark 27, we package the pair $(\text{in}_p, \text{out}_p)$ as a single isomorphism in_p .

The presheaf of terms in the model \mathbf{A} internalizes as a meta-space type of variables which by virtue of the structure map $\mathbf{A} \rightarrow \mathbf{I}$ can be indexed over the object-space collection of terms. We realize this synthetically as follows:

$$\text{var} : (A : \text{tp}) \rightarrow \{\mathcal{U} | \text{obj} \hookrightarrow \text{tm} A\}$$

We refer to extensional type theory extended with these constants and modalities as the language of STC.

Remark 30. To account for strict universes – those for which el commutes strictly with chosen codes – some prior STC developments employed strict gluing along the image of el (Sterling 2021; Sterling and Angiuli 2021). By limiting our usage of strict gluing to obj , we are able to execute our constructions in a constructive metatheory. To model strict universes, we instead use the cumulativity of the hierarchy of universes \mathcal{U}_i and the fact that all levels are coherently closed under modalities and strict gluing.

6.6 Normal and neutral forms

Internally to STC, we now specify the normal and neutral forms of terms, and the normal forms of types. Following Sterling and Angiuli (2021), we index the type of neutral forms by a *frontier of instability*, a proposition at which the neutral form is no longer meaningful. Our construction

$$\begin{aligned}
 \text{var} &: (x : \text{tm } A) \rightarrow \text{ne}_\bullet A \perp x \\
 \text{unst} &: \{a : \text{tm } A\} \rightarrow \text{ne}_\bullet A \top a \\
 _ &: (a : \text{tm } A) (e : \text{ne}_\bullet A \top a) \rightarrow e = \text{unst } a \\
 \\
 \text{ext}_p &: \text{nftp}_\bullet A \rightarrow (\{p\} \text{nf}_\bullet A a) \rightarrow \text{nftp}_\bullet (\text{ext}_p A a) \\
 \text{in}_p &: \text{nftp}_\bullet A \rightarrow \text{nf}_\bullet A u \rightarrow \text{nf}_\bullet (\text{ext}_p A a) u \\
 \text{out}_p &: \text{nftp}_\bullet A \rightarrow \text{ne}_\bullet (\text{ext}_p A a) \phi u \rightarrow \text{ne}_\bullet A (\phi \vee \langle p \rangle) (\text{in}_p^{-1} u) \\
 \\
 \text{up}_{\text{uni}} &: \text{ne}_\bullet \text{uni } \phi A \rightarrow (\{\phi\} \text{nftp}_\bullet (\text{el } A)) \rightarrow \text{nftp}_\bullet (\text{el } A) \\
 \text{up}_{\text{el}} &: \text{ne}_\bullet \text{uni } \phi A \rightarrow (\{\phi\} \text{nftp}_\bullet (\text{el } A)) \rightarrow \text{ne}_\bullet (\text{el } A) \psi a \rightarrow (\{\phi \vee \psi\} \text{nf}_\bullet (\text{el } A) a) \rightarrow \\
 &\quad \text{nf}_\bullet (\text{el } A) a \\
 \text{el} &: \text{ne}_\bullet \text{uni } \phi A \rightarrow (\{\phi\} \text{nf}_\bullet \text{uni } A) \rightarrow \text{nf}_\bullet \text{uni } A \\
 \\
 _ &: (e : \text{ne}_\bullet \text{uni } \top A) (u : \text{nftp}_\bullet A) \rightarrow \text{up}_{\text{uni}} e u = u \\
 _ &: (e : \text{ne}_\bullet \text{uni } \top A) (u : \text{nf}_\bullet \text{uni } A) \rightarrow \text{el } e u = u \\
 _ &: \{\phi \vee \psi\} (e_A : \text{ne}_\bullet \text{uni } \phi A) (u_A : \{\phi\} \text{nftp}_\bullet (\text{el } A)) (e_a : \text{ne}_\bullet (\text{el } A) \psi a) (u_a : \text{nf}_\bullet A a) \\
 &\quad \rightarrow \text{up}_{\text{el}} e_A u_A e_a u_a = u_a
 \end{aligned}$$

Figure 2. Selected rules from the definition of nf_\bullet , ne_\bullet , and nftp_\bullet .

proceeds in two steps. First, we define a series of indexed quotient-inductive definitions (Kaposi et al. 2019) specifying the meta-space components of normal and neutral forms:

$$\begin{aligned}
 \text{nf}_\bullet &: (A : \text{tp}) \rightarrow \text{tm } A \rightarrow \mathcal{U}_\bullet^0 \\
 \text{ne}_\bullet &: (A : \text{tp}) \rightarrow \Omega_{\text{dec}} \rightarrow \text{tm } A \rightarrow \mathcal{U}_\bullet^0 \\
 \text{nftp}_\bullet &: \text{tp} \rightarrow \mathcal{U}_\bullet^0
 \end{aligned}$$

Next, we use the strict gluing connective to define the types of normals, neutrals, and normal types such that they lie strictly over tm and tp :

$$\begin{aligned}
 \text{nf } A &= (a : \text{tm } A) \ltimes \text{nf}_\bullet A a \\
 \text{ne}_\phi A &= (a : \text{tm } A) \ltimes \text{ne}_\bullet A \phi a \\
 \text{nftp} &= (A : \text{tp}) \ltimes \text{nftp}_\bullet A
 \end{aligned}$$

We illustrate a representative fragment of the inductive definitions in Figure 2.

The induction principles for nf_\bullet , ne_\bullet and nftp_\bullet play no role in the main development, which works with any algebra for these constants. These induction principles, however, are needed in order to prove Theorem 48 and deduce the decidability of definitional equality and the injectivity of type constructors. These same considerations motivate our choice to index ne_\bullet over Ω_{dec} rather than Ω .

6.7 A glued normalization algebra

We can now construct a new $\mathbf{TT}_{\mathbb{P}}$ -algebra internally to $\mathbf{Sh } \mathbf{G}$, satisfying the constraint that each of its constituents restricts under obj to the corresponding constant from the $\mathbf{TT}_{\mathbb{P}}$ -algebra inherited from $\mathbf{Sh } \mathbf{I}$. We shall refer to this as the *normalization algebra*. For instance, we must define types representing object types and terms:

$$\text{tp}^* : \{\mathcal{U}_2 | \text{obj} \hookrightarrow \text{tp}\} \quad \text{tm}^* : \{\text{tp}^* \rightarrow \mathcal{U}_1 | \text{obj} \hookrightarrow \text{tm}\}$$

The meta-space component of the computability structure of types is given as a dependent record below:

record $\text{tp}_\bullet (A : \text{tp}) : \mathcal{U}_2$ **where**

code : $\text{nftp}_\bullet A$

$\text{tm}_\bullet : \text{tm}A \rightarrow \mathcal{U}^1_\bullet$

reflect : $(a : \text{tm}A) (\phi : \Omega_{\text{dec}}) (e : \text{ne}_\bullet A \phi a) \rightarrow (a_\phi : \{\phi\} \text{tm}_\bullet a) \rightarrow \{\text{tm}_\bullet a | \phi \hookrightarrow a_\phi\}$

reify : $(a : \text{tm}A) \rightarrow \text{tm}_\bullet a \rightarrow \text{nf}_\bullet A a$

The tm_\bullet field classifies the meta-space component of a given element; the reflect and reify fields generalize the familiar operations of normalization by evaluation, subject to Sterling and Angiuli’s stabilization yoga (Sterling and Angiuli 2021). We finally define both tp and tm using strict gluing to achieve the correct boundary:

$\text{tp}^* = (A : \text{tp}) \ltimes \text{tp}_\bullet A$

$\text{tm}^* A = (a : \text{tm}A) \ltimes (\text{ungl}A).\text{tm}_\bullet a$

Notation 31. Henceforth, we will write $A.\text{fld}$ rather than $(\text{ungl}A).\text{fld}$ to access a field of the closed component of A .

We must also define $\langle p \rangle^* : \Omega_{\text{dec}}$ for each $p \in \mathbb{P}$ subject to the condition that obj implies $\langle p \rangle^* = \langle p \rangle$. As there is no normalization data associated with these propositions, we define $\langle p \rangle^* = \langle p \rangle$ which clearly satisfies the boundary condition. It remains to show that $(\text{tp}^*, \text{tm}^*)$ are closed under all the connectives of $\mathbf{TT}_{\mathbb{P}}$. We show two representative cases: extension types and the universe.

6.7.1 Extension types

Fixing $A : \text{tp}^*, p : \mathbb{P}, a : \{\langle p \rangle\} \text{tm}^* A$, we must construct the following pair of constants:

$\text{ext}_p^* A a : \{\text{tp}^* | \text{obj} \hookrightarrow \text{ext}_p^* A a\}$

$\text{in}_p^* A a : \{\{\text{tm}^* A | \langle p \rangle \hookrightarrow a\} \cong \text{tm}^* (\text{ext}_p^* A a) | \text{obj} \hookrightarrow \text{in}_p^* A a\}$

Recalling the definition of tp as a strict gluing type, we observe that the boundary condition on ext_p already fully constrains the first component:

$\text{ext}_p A a = [\text{obj} \hookrightarrow \text{ext}_p A a \mid \text{?} : \text{tp}_\bullet (\text{ext}_p A a)]$

In the above, we have used Notation 28 for constructing elements of a strict gluing type.

We define the second component as follows, using copattern matching notation:

$(\text{ext}_p A a).\text{code} = \text{ext}_p A a.\text{code} (A.\text{reify } a)$

$(\text{ext}_p A a).\text{tm}_\bullet x = \bullet \{A.\text{tm}_\bullet (\text{in}_p^{-1} x) \} | \langle p \rangle \hookrightarrow a$

$(\text{ext}_p A a).\text{reify } (\eta_\bullet x) = \text{in}_p A a.\text{code} (A.\text{reify } x)$

$(\text{ext}_p A a).\text{reflect } x \phi e (\eta_\bullet x_\phi) =$

$\eta_\bullet \left(\begin{array}{l} A.\text{reflect} \\ (\text{in}_p^{-1} x) (\phi \vee \langle p \rangle) \\ (\text{out}_p A a.\text{code } e) [\phi \hookrightarrow x_\phi | \langle p \rangle \hookrightarrow a] \end{array} \right)$

In the clauses of reify and reflect, we were allowed to assume that the argument was of the form $\eta_\bullet x$ where η_\bullet is the unit of the modality $\bullet : \eta_\bullet : A \rightarrow \bullet A$. This is because we are mapping into meta-space types and so this “pattern-matching” amounts to the bind operation of the monad \bullet .

record $\text{uni}_\bullet : \mathcal{U}_1$ **where**
 $\text{code} : \text{nf}_\bullet \text{ uni } A$
 $\text{el-code} : \text{nftp}_\bullet (\text{el } A)$
 $\text{tm}_\bullet : \text{tm} (\text{el } A) \rightarrow \mathcal{U}_\bullet^0$
 $\text{reflect} : (a : \text{tm } A) (\phi : \Omega_{\text{dec}}) \rightarrow (e : \text{ne}_\bullet A \phi a) \rightarrow (a_\phi : \{\phi\} \text{tm}_\bullet a) \rightarrow \{\text{tm}_\bullet a \mid \phi \hookrightarrow a_\phi\}$
 $\text{reify} : (a : \text{tm } A) \rightarrow \text{tm}_\bullet a \rightarrow \text{nf}_\bullet A a$

$\text{uni}^* : \{\text{tp}^* \mid \text{obj} \hookrightarrow \text{uni}\}$
 $\text{uni}^*.\text{code} = \text{uni}$
 $\text{uni}^*.\text{tm}_\bullet A = \bullet (\text{uni}_\bullet A)$
 $\text{uni}^*.\text{reify } _ (\eta_\bullet A) = A.\text{code}$

$\text{obj} \hookrightarrow \text{uni}^*.\text{reflect } A \phi e_A A \phi = A$
 $\text{ungl} (\text{uni}^*.\text{reflect } A \phi e_A A \phi) =$
let $A_\phi : \{\phi\} \text{uni}_\bullet A = X \leftarrow A_\phi; X;$
 η_\bullet record
 $\text{code} = \text{up}_{\text{uni}} e_A A_\phi.\text{code}$
 $\text{el-code} = \text{el } e_A A_\phi.\text{el-code}$
 $\text{tm}_\bullet a = \bullet ((u : \text{nf}_\bullet A x) \times \{\phi\} \{a : A_\phi.\text{tm}_\bullet a \mid A_\phi.\text{reify } a = u\})$
 $\text{reflect } a \psi e_a a_\psi =$
let $a_\psi : \{\psi\} (u : \text{nf}_\bullet A x) \times \dots = x \leftarrow a_\psi; x;$
let $a_\phi = A_\phi.\text{reflect } \psi e_a a_\psi;$
 $\eta_\bullet (\text{el } e_A A_\phi.\text{el-code } e_a [\phi \hookrightarrow A_\phi.\text{reify } a_\phi \mid \psi \hookrightarrow \pi_1 a_\psi], a_\phi)$
 $\text{reify } _ (\eta_\bullet (u, _)) = u$

Figure 3. The normalization structure on the universe.

Remark 32. Stabilized neutrals are crucial to the definition of $(\text{ext}_p^* A a).\text{reflect}$ above: without them, we could not ensure that reflecting $\text{out}_p A.\text{code } e$ lies within the specified subtype of $A.\text{tm}_\bullet$.

The definition of in_p is now straightforward:

$$\text{in}_p^* A a x = [\text{obj} \hookrightarrow \text{in}_p A a x | \text{ungl } x]$$

We leave the routine verification of the various boundary conditions to the reader; nearly all of them follow immediately from the properties of strict gluing.

6.7.2 The universe

We now turn to the construction of the universe in the normalization algebra; it is here that the complexity of unstable neutrals becomes evident. Once again the boundary conditions on uni force part of its definition:

$$\text{uni}^* = [\text{obj} \hookrightarrow \text{uni} \mid \text{=?} : \text{tm}_\bullet \text{ uni}]$$

The second component of uni^* is complex, and we present its definition in Figure 3. The inclusion of el-code in uni_\bullet is necessary in order to define el^* :

$$\begin{aligned} \text{obj} \hookrightarrow \text{el}^* A &= \text{el } A \\ (\text{el}^* (\eta_\bullet A)).\text{code} &= A.\text{el-code} \\ (\text{el}^* (\eta_\bullet A)).\text{tm}_\bullet &= A.\text{tm}_\bullet \\ (\text{el}^* (\eta_\bullet A)).\text{reflect} &= A.\text{reflect} \\ (\text{el}^* (\eta_\bullet A)).\text{reify} &= A.\text{reify} \end{aligned}$$

Finally, we must show that uni is closed under all small type formers and that el preserves them. This flows from the cumulativity of universes in $\mathbf{Sh\ G}$; to close uni under, for example, products, we essentially "redo" the construction of products in \mathbf{tp} by altering its predicate to be valued in \mathcal{U}_0 rather than \mathcal{U}_1 .

6.7.3 The evaluation functor

In this section, we equip $\mathbf{Sh\ G}$ with the maximal CwR structure, in which all maps are representable.⁵ We have just now defined an interpretation of $\mathbf{TT}_{\mathbb{P}}$'s signature (Section 6.2) in $\mathbf{Sh\ G}$, and so by the universal property (Proposition 10) of \mathbb{T} as the classifying CwR for this signature, we obtain a unique CwR functor $I_{\mathbf{Sh\ G}} : \mathbb{T} \rightarrow \mathbf{Sh\ G}$ sending every construct of $\mathbf{TT}_{\mathbb{P}}$ to its interpretation.

6.8 The normalization algorithm

Having constructed the normalization algebra in $\mathbf{Sh\ G}$ (Section 6.7), we can now define the actual normalization function using an argument based on those presented by Fiore (2002, 2022, §II.2) and Sterling (2025, §3.3), making use of the inserter model of atomic substitutions introduced by Bocquet *et al.* (2021). As our results are constructive, our normalization function corresponds to an actual normalization by evaluation algorithm – whose executable computational content Fiore (2022) has demonstrated explicitly in the simply typed case.

6.8.1 Stripping of atomic contexts

We first must establish an intermediate result: that the functor $\mathbf{A}(-) : \mathbb{P} \rightarrow \mathbf{A}_{\diamond}$ sending each $p \in \mathbb{P}$ to the associated unary atomic context is fully faithful and has a left adjoint, that is, that \mathbb{P} is reflective in \mathbf{A}_{\diamond} . The left adjoint allows an atomic context to be "stripped" of anything that induces variables, leaving only propositional assumptions. This result is ultimately used in Lemma 40 to exhibit an isomorphism $\mathbf{A}\langle p \rangle \cong \alpha^* \mathbf{I}\langle p \rangle$.

While it is straightforward to imagine how such a reflection can be defined by "induction on atomic contexts and repeated weakening," we have not given an inductive specification of \mathbf{A}_{\diamond} and instead opted to specify it through its universal property. Accordingly, we define this stripping map using a model to which we may apply the universal property of \mathbf{A} . Fundamentally, however, the resulting constructions are the same, but our insistence on using only these universal properties enables us to avoid fixing a particular and explicit construction of atomic contexts.

Construction 33 (*The stripping model*). We consider a model \mathbf{P} of \mathbb{T} in which we set $\mathbf{P}_{\diamond} = \mathbb{P}$, $\mathbf{P}(\text{tp}) = \mathbf{P}(\text{tm}) = \mathbf{I}_{\mathbf{Pr}\ \mathbb{P}}$, and $\mathbf{P}\langle p \rangle = \vee_{\mathbb{P}} p$. All the remaining constructs of the model are trivial by virtue of these definitions. From the universal property of \mathbf{I} as the bi-initial \mathbb{T} -model, we obtain a unique homomorphism of models $I_{\mathbf{P}} : \mathbf{I} \rightarrow \mathbf{P}$ whose contextual component is a product preserving functor $I_{\mathbf{P}}^{\diamond} : \mathbf{I}_{\diamond} \rightarrow \mathbb{P}$.

Lemma 34. *The component $\mathbf{A}(-) : \mathbb{P} \rightarrow \mathbf{A}_{\diamond}$ of the bi-initial atomic substitution model over \mathbf{I} is full and faithful.*

Proof. Any functor out of a poset is necessarily faithful. To see that $\mathbf{A}(-) : \mathbb{P} \rightarrow \mathbf{A}_{\diamond}$ is full, we fix a morphism $\mathbf{A}\langle p \rangle \rightarrow \mathbf{A}\langle q \rangle$ in \mathbf{A}_{\diamond} ; as this morphism is necessarily unique, it suffices to show that $p \leq q$ in \mathbb{P} . We consider the image of $\mathbf{A}\langle p \rangle \rightarrow \mathbf{A}\langle q \rangle$ under the contextual component of the composite homomorphism $I_{\mathbf{P}} \circ \alpha : \mathbf{A} \rightarrow \mathbf{P}$ of \mathbb{T}_0 -models, which gives precisely the desired inequality $p \leq q$, recalling that each $\langle r \rangle$ is representable in \mathbb{T}_0 and thus preserved by homomorphisms of models. \square

Lemma 35. *The functor between categories of contexts induced by $I_{\mathbf{P}} \circ \alpha : \mathbf{A} \rightarrow \mathbf{P}$ is left adjoint to the embedding $\mathbf{A}(-) : \mathbf{P} \hookrightarrow \mathbf{A}_{\diamond}$.*

Proof. The counit in \mathbb{P} is given by the identity inequality, as each $\langle p \rangle$ is representable in \mathbb{T}_0 and thus preserved by homomorphisms. For the unit, we must construct a (necessarily unique) arrow $\Gamma \rightarrow \mathbf{A}\langle I_{\mathbf{p}}^{\diamond}(\alpha_{\diamond}\Gamma) \rangle$ in \mathbf{A}_{\diamond} for each atomic context Γ .

For this, we consider a new atomic substitution model \mathbf{E} over \mathbf{A} whose category of contexts \mathbf{E}_0 is the following inserter object (Lack 2009, Section 6.5) in \mathbf{Cat} :

$$\mathbf{E}_\diamond \overset{\phi_\diamond}{\dashrightarrow} \mathbf{A}_\diamond \overset{1_{\mathbf{A}_\diamond}}{\underset{\mathbf{A}\langle-\rangle \circ I_{\mathbf{p}}^\diamond \circ \alpha_\diamond}{\rightrightarrows}} \mathbf{A}_\diamond$$

Equivalently, \mathbf{E}_\diamond is the full subcategory of \mathbf{A}_\diamond spanned by atomic contexts Γ for which there exists an arrow $\Gamma \rightarrow \mathbf{A}(\mathbf{I}_\mathbf{P}^\diamond(\alpha_\diamond \Gamma))$; as the codomain is subterminal, such arrows are necessarily unique. We define all the constructs of \mathbb{T}_0 in \mathbf{E} as in \mathbf{A} , and it remains only to check that \mathbf{E} has a terminal object and is closed under context comprehension and phase comprehension.

- (1) For the terminal object, we see that $\mathbf{A}(I_{\mathbf{P}}^{\diamond}(\alpha_{\diamond}))$ is already terminal.
- (2) For the context comprehension, we fix $\Gamma \in \mathbf{E}_{\diamond}$ and $A \in \mathbf{A}(\mathrm{tp}(\Gamma))$, and we must check that there exists a map $\Gamma.A \rightarrow \mathbf{A}(I_{\mathbf{P}}^{\diamond}(\alpha_{\diamond}(\Gamma.A)))$. As $\alpha \circ I_{\mathbf{P}}$ is a homomorphism of models, it preserves context comprehensions; unraveling definitions, we ultimately have $I_{\mathbf{P}}^{\diamond}(\alpha_{\diamond}(\Gamma.A)) = I_{\mathbf{P}}^{\diamond}(\alpha_{\diamond}(\Gamma))$ and so we are done.
- (3) For phase comprehension, we fix $\Gamma \in \mathbf{E}_{\diamond}$ and $p \in \mathbb{P}$ to check that there exists an arrow $\Gamma.\mathbf{A}\langle p \rangle \rightarrow \mathbf{A}(I_{\mathbf{P}}^{\diamond}(\alpha_{\diamond}(\Gamma.\mathbf{A}\langle p \rangle)))$. But we have $I_{\mathbf{P}}^{\diamond}(\alpha_{\diamond}(\Gamma.\mathbf{A}\langle p \rangle)) = I_{\mathbf{P}}^{\diamond}(\alpha_{\diamond}(\Gamma)) \wedge p$, so we may use the projection $\Gamma.\mathbf{A}\langle p \rangle \rightarrow \mathbf{A}\langle p \rangle$.

We evidently have a homomorphism of \mathbb{T}_0 -models $\eta : \mathbf{E} \rightarrow \mathbf{A}$ that exhibits \mathbf{E} as a atomic substitution model over \mathbf{A} . Postcomposing with the structure map $\alpha : \mathbf{A} \rightarrow \mathbf{I}$, we can view \mathbf{E} as a atomic substitution model over \mathbf{I} . Thus, by the universal property of \mathbf{A} , we have a universal section $J_{\mathbf{E}} : \mathbf{A} \rightarrow \mathbf{E}$ to $\eta : \mathbf{E} \rightarrow \mathbf{A}$. This shows that every atomic context $\Gamma \in \mathbf{A}_{\diamond}$ can be equipped with an arrow $\Gamma \rightarrow \mathbf{A}(I_{\mathbf{P}}^{\diamond}(\alpha_{\diamond}\Gamma))$. Assembling all these arrows together, we have the unit of the adjunction $I_{\mathbf{P}}^{\diamond} \circ \alpha_{\diamond} \dashv \mathbf{A}(-)$. \square

The force of Lemma 35 is to show that \mathbb{P} is a reflective subcategory of \mathbf{A}_\diamond .

6.8.2 Computability spaces of atomic and computable substitutions

We will consider two computability spaces induced by an atomic context Γ : the computability space $\llbracket \Gamma \rrbracket$ of “computable substitutions into Γ ” and the computability space $\langle \Gamma \rangle$ of “atomic substitutions into Γ .”

Construction 36 (*The computability space of computable substitutions*). The computability space $\llbracket \Gamma \rrbracket$ of computable substitutions into an atomic context Γ is defined in terms of the interpretation of \mathbb{T} into **Sh G** as follows, sending each atomic context to the computability space determined by the algebra structure:

$$\begin{array}{ccccc} \mathbf{A}_\diamond & \xrightarrow{\alpha_\diamond} & \mathbf{I}_\diamond & \xhookrightarrow{\subseteq} & \mathbb{T} \xrightarrow{I_{\text{Sh } G}} \mathbf{Sh } G \\ & \searrow & & & \uparrow \\ & & \llbracket - \rrbracket & & \end{array}$$

Construction 37 (*The computability space of atomic substitutions*). We define an embedding $\langle \!| - \!| \rangle : \mathbf{A}_\diamond \hookrightarrow \mathbf{Sh} \mathbf{G}$ sending $\Gamma \in \mathbf{A}_\diamond$ to the computability space $\langle \!| \Gamma \!| \rangle$ with $j^*(\langle \!| \Gamma \!| \rangle) = \mathbf{YI}_\diamond \alpha_\diamond \Gamma$ and $i^*(\langle \!| \Gamma \!| \rangle) = \mathbf{YA}_\diamond \Gamma$, such that $\pi_{\langle \!| \Gamma \!| \rangle} \rightarrow \mathbf{YA}_\diamond \Gamma \xrightarrow{\alpha^*} \mathbf{YI}_\diamond \alpha_\diamond \Gamma$ is defined on generalized elements by the functorial action of $\alpha_\diamond : \mathbf{A}_\diamond \rightarrow \mathbf{I}_\diamond$ as follows:

$$\pi_{\langle \!| \Gamma \!| \rangle}^\Delta(\gamma : \Delta \rightarrow \Gamma) = \alpha_\diamond \gamma : \alpha_\diamond \Delta \rightarrow \alpha_\diamond \Gamma$$

Miraculously, the coreflective embedding $i_! : \mathbf{Sh} \mathbf{A} \rightarrow \mathbf{Sh} \mathbf{G}$ sends $\mathbf{YA}_\diamond \Gamma$ to *precisely* the computability space $\langle \!| \Gamma \!| \rangle$, up to isomorphism:

$$\begin{aligned} i_! \mathbf{YA}_\diamond \Gamma &= (\alpha_! \mathbf{YA}_\diamond \Gamma, : \mathbf{YA}_\diamond \Gamma \rightarrow \alpha^* \alpha_! \mathbf{YA}_\diamond \Gamma) \\ &\cong (\mathbf{YI}_\diamond \alpha_\diamond \Gamma, : \mathbf{YA}_\diamond \Gamma \rightarrow \alpha^* \mathbf{YI}_\diamond \alpha_\diamond \Gamma) \\ &= \langle \!| \Gamma \!| \rangle \end{aligned}$$

Note that the functors $\llbracket - \rrbracket, \langle \!| - \!| \rangle$ lift into the slice $\mathbf{Cat} \downarrow \mathbf{Sh} \mathbf{I}$ in the following sense:

$$\begin{array}{ccccc} \mathbf{Sh} \mathbf{G} & \xleftarrow{\langle \!| - \!| \rangle} & \mathbf{A}_\diamond & \xrightarrow{\llbracket - \rrbracket} & \mathbf{Sh} \mathbf{G} \\ & \searrow \scriptstyle \hookrightarrow^* & \downarrow \scriptstyle \mathbf{YI}_\diamond \circ \alpha_\diamond & \swarrow \scriptstyle \hookrightarrow^* & \\ & & \mathbf{Sh} \mathbf{I} & & \end{array}$$

Notation 38. Let Γ be an atomic context, and let $A : \langle \!| \Gamma \!| \rangle \rightarrow \mathbf{tp}$ be an object-space type, which we may regard as a morphism $\alpha_\diamond \Gamma \rightarrow \mathbf{tp}$ in \mathbb{T} . We shall write $\llbracket A \rrbracket : \llbracket \Gamma \rrbracket \rightarrow \mathbf{tp}^*$ for the image of A under the interpretation functor $I_{\mathbf{Sh} \mathbf{G}}$.

Lemma 39. Let Γ be an atomic context, and let $A : \langle \!| \Gamma \!| \rangle \rightarrow \mathbf{tp}$ be an object-space type (which we may regard as a morphism $\alpha_\diamond \Gamma \rightarrow \mathbf{tp}$ in \mathbb{T}). Then we have the following cartesian squares:

$$\begin{array}{ccc} \langle \!| \Gamma.A \!| \rangle \rightarrow \sum_{A:\mathbf{tp}} \mathbf{var} A & & \llbracket \Gamma.A \rrbracket \rightarrow \sum_{A:\mathbf{tp}^*} \mathbf{tm}^* A \\ \downarrow \scriptstyle \langle \!| p_A \!| \rangle & \downarrow \scriptstyle \pi_1 & \downarrow \scriptstyle \pi_1 \\ \langle \!| \Gamma \!| \rangle \xrightarrow{A} \mathbf{tp} & & \llbracket \Gamma \rrbracket \xrightarrow{\llbracket A \rrbracket} \mathbf{tp}^* \end{array}$$

Stated in the internal language, we have canonical isomorphisms $\langle \!| \Gamma.A \!| \rangle \cong \sum_{\gamma:\langle \!| \Gamma \!| \rangle} \mathbf{var} (A\gamma)$ and $\llbracket \Gamma.A \rrbracket \cong \sum_{\gamma:\llbracket \Gamma \rrbracket} \mathbf{tm} (\llbracket A \rrbracket \gamma)$.

Proof. The latter is the image of a pullback square in \mathbb{T} under $I_{\mathbf{Sh} \mathbf{G}}$, which is finitely continuous. The former can be seen by means of an explicit computation. \square

We now come to an important result relating the interpretation of $\langle p \rangle$ in \mathbf{I} to the interpretation of the same in \mathbf{A} . Lemma 40 is the *raison d'être* for the stripping model \mathbf{P} in Section 6.8.1.

Lemma 40. We have a (necessarily unique) isomorphism $\mathbf{A}\langle p \rangle \cong \alpha^* \mathbf{I}\langle p \rangle$.

Proof. As both presheaves are subterminal, it is enough to see that one is inhabited if and only if the other is.

- (1) We may transpose a map $y_{A_\diamond} \Gamma \rightarrow \alpha^* \mathbf{I}\langle p \rangle$ to get $\alpha_\diamond \Gamma \rightarrow \mathbf{I}\langle p \rangle$; applying the functorial action of $I_p^\diamond \rightarrow \mathbf{I}_\diamond \mathbb{P}$, we have $I_p^\diamond \alpha_\diamond \Gamma \leq I_p^\diamond \mathbf{I}\langle p \rangle = p$ and by adjoint transpose with Lemma 35, we have $\Gamma \rightarrow \mathbf{A}\langle p \rangle$.
- (2) Conversely, given an arrow $\Gamma \rightarrow \mathbf{A}\langle p \rangle$ we may apply the functorial action of $\alpha_\diamond : \mathbf{A}_\diamond \rightarrow \mathbf{I}_\diamond$ to obtain an arrow $\alpha_\diamond \Gamma \rightarrow \alpha_\diamond \mathbf{A}\langle p \rangle$. As $\langle p \rangle$ is representable in \mathbb{T}_0 , it is preserved by morphisms of models like $\alpha : \mathbf{A} \rightarrow \mathbf{I}$; thus, $\alpha_\diamond \mathbf{A}\langle p \rangle \cong \mathbf{I}\langle p \rangle$ and so we have $\alpha_\diamond \Gamma \rightarrow \mathbf{I}\langle p \rangle$, which we may transpose to obtain $y_{A_\diamond} \Gamma \rightarrow \alpha^* \mathbf{I}\langle p \rangle$. \square

6.8.3 Hydration of atomic substitutions

A critical point in concrete normalization by evaluation algorithms is to “reflect” a vector of variables as an environment of (computable) values against which the computability interpretation of an open term can be executed; in a concrete setting, this operation is defined by recursion on the atomic contexts. The same process, which we shall refer to here as the *hydration of atomic substitutions*, plays an equally important role in semantic proofs of normalization in the guise of a certain *hydration map* $\nearrow : \langle - \rangle \mapsto \llbracket - \rrbracket$ in $\text{Cat} \downarrow \mathbf{Sh} \mathbf{I}$ that we shall need to construct.

Just as in the definition of the stripping map, we are confronted by the fact that we have defined the atomic contexts only by means of a universal property (the bi-initial atomic substitution model over \mathbf{I}), so we do not immediately have anything concrete to do recursion on. The innovation of Bocquet et al. (2021) was to find the correct categorical induction motive that explains the usual recursive argument purely in terms of the universal property of the bi-initial atomic substitution model (likewise due to op. cit.). In what follows, we adapt their ideas to our setting and show how to construct the desired hydration map.

We begin by defining a new model \mathbf{H} of \mathbb{T}_0 that we shall refer to as the *hydration model*. The category of contexts \mathbf{H}_\diamond is defined to be the underlying category of the inserter object for $\langle - \rangle, \llbracket - \rrbracket : \mathbf{A}_\diamond \rightarrow \mathbf{Sh} \mathbf{G}$ in $\text{Cat} \downarrow \mathbf{Sh} \mathbf{I}$ as depicted below:

$$\begin{array}{ccccc}
 \mathbf{H}_\diamond & \xrightarrow{\psi_\diamond} & \mathbf{A}_\diamond & \xrightarrow{\langle - \rangle} & \mathbf{Sh} \mathbf{G} \\
 & \searrow & \downarrow y_{\mathbf{I}_\diamond} \circ \alpha_\diamond & \searrow \llbracket - \rrbracket & \\
 & & \mathbf{Sh} \mathbf{I} & \xleftarrow{j^*} &
 \end{array}$$

Explicitly, an object of \mathbf{H}_\diamond is a pair (Γ, h_Γ) of an object $\Gamma \in \mathbf{A}_\diamond$ together with an arrow $h_\Gamma : \langle \Gamma \rangle \rightarrow \llbracket \Gamma \rrbracket$ whose image under $j^* : \mathbf{Sh} \mathbf{G} \rightarrow \mathbf{Sh} \mathbf{I}$ is the identity map on $y_{\mathbf{I}_\diamond} \alpha_\diamond \Gamma$. An arrow from (Δ, h_Δ) to (Γ, h_Γ) is given by an arrow $\gamma : \Delta \rightarrow \Gamma$ in \mathbf{A}_\diamond making the following square commute:

$$\begin{array}{ccc}
 \langle \Delta \rangle & \xrightarrow{\langle \gamma \rangle} & \langle \Gamma \rangle \\
 h_\Delta \downarrow & & \downarrow h_\Gamma \\
 \llbracket \Delta \rrbracket & \xrightarrow{\llbracket \gamma \rrbracket} & \llbracket \Gamma \rrbracket
 \end{array}$$

Clearly, \mathbf{H}_\diamond has a terminal object because $\llbracket \mathbf{I}_{A_\diamond} \rrbracket$ is terminal. Anticipating that ψ_\diamond should lift to a morphism of models exhibiting \mathbf{H} as a atomic substitution model over \mathbf{A} , we define $\mathbf{H}\langle p \rangle = \psi_\diamond^* \mathbf{A}\langle p \rangle$ and $\mathbf{H}(\text{tp}) = \psi_\diamond^* \mathbf{A}(\text{tp})$. In order to define $\mathbf{H}(\text{tm})$, it will be easiest to first define context comprehensions.

Construction 41 (*Context comprehensions in \mathbf{H}_\diamond*). Given a context (Γ, h_Γ) in \mathbf{H}_\diamond and a type $A : \mathbf{Y}_{\mathbf{A}_\diamond} \Gamma \rightarrow \mathbf{A}(\text{tp})$, we can lift the context comprehension $\Gamma.A \in \mathbf{A}_\diamond$ into \mathbf{H}_\diamond by finding a suitable map $h_{\Gamma.A} : \langle \Gamma.A \rangle \rightarrow \llbracket \Gamma.A \rrbracket$ whose image in $\mathbf{Sh} \mathbf{I}$ is the identity. Recalling Lemma 39, we may equivalently construct (from the internal point of view) the following map:

$$h_{\Gamma.A} : \{ (\Sigma_{\gamma : \langle \Gamma \rangle} \text{var } (A\gamma)) \rightarrow \Sigma_{\gamma : \llbracket \Gamma \rrbracket} \text{tm}(\llbracket A \rrbracket \gamma) | \text{obj} \hookrightarrow \lambda u. u \}$$

$$h_{\Gamma.A}(\gamma, x) = h_{\Gamma} \gamma, \llbracket A \rrbracket(h_{\Gamma} \gamma). \text{reflect } x \perp (\text{var } x) (\lambda_- : \perp. \star)$$

The projection $p_A : \Gamma.A \rightarrow \Gamma$ tracks a morphism $(\Gamma.A, h_{\Gamma.A}) \rightarrow (\Gamma, h_\Gamma)$, by definition of $h_{\Gamma.A}$.

Construction 42 (*Phase comprehensions in \mathbf{H}_\diamond*). Given $\Gamma \in \mathbf{H}_\diamond$ and $p \in \mathbb{P}$, we must exhibit a map $h_{\Gamma.A\langle p \rangle} : \langle \Gamma.A\langle p \rangle \rangle \rightarrow \llbracket \Gamma.A\langle p \rangle \rrbracket$. Using Lemma 39, we construct this by combining the assumed map $h_\Gamma : \langle \Gamma \rangle \rightarrow \llbracket \Gamma \rrbracket$ with the (necessarily unique) map $\langle A\langle p \rangle \rangle \rightarrow \llbracket A\langle p \rangle \rrbracket$ obtained from the identity map on $A\langle p \rangle$ in the following way:

(1) First, we observe that $\llbracket A\langle p \rangle \rrbracket \cong \mathbf{j}_* \mathbf{I}\langle p \rangle$ as follows:

$$\begin{aligned} \llbracket A\langle p \rangle \rrbracket &= I_{\mathbf{Sh} \mathbf{G}} \alpha_\diamond A\langle p \rangle && \text{by definition} \\ &\cong I_{\mathbf{Sh} \mathbf{G}} \mathbf{I}\langle p \rangle && \alpha \text{ is a homomorphism} \\ &\cong \mathbf{j}_* \mathbf{I}\langle p \rangle && \text{by definition} \end{aligned}$$

(2) Then we proceed by adjoint calisthenics:

$$\begin{aligned} \text{hom}_{\mathbf{Sh} \mathbf{G}} (\langle A\langle p \rangle \rangle \llbracket A\langle p \rangle \rrbracket) &\cong \text{hom}_{\mathbf{Sh} \mathbf{G}} (\langle A\langle p \rangle \rangle, \mathbf{j}_* \mathbf{I}\langle p \rangle) && \text{by the above} \\ &\cong \text{hom}_{\mathbf{Sh} \mathbf{I}} (\mathbf{j}^* \langle A\langle p \rangle \rangle, \mathbf{I}\langle p \rangle) && \text{by adjoint transpose} \\ &\cong \text{hom}_{\mathbf{Sh} \mathbf{I}} (\alpha! A\langle p \rangle, \mathbf{I}\langle p \rangle) && \text{by definition of } \langle - \rangle \\ &\cong \text{hom}_{\mathbf{Sh} \mathbf{A}} (A\langle p \rangle, \alpha^* \mathbf{I}\langle p \rangle) && \text{by adjoint transpose} \\ &\cong \text{hom}_{\mathbf{Sh} \mathbf{A}} (A\langle p \rangle, A\langle p \rangle) && \text{by Lemma 40} \end{aligned}$$

Construction 43 (*The presheaf of terms*). We define $\mathbf{H}(\text{tm}) \in \mathbf{Pr} \mathbf{H}_\diamond \downarrow \mathbf{H}(\text{tp})$ to send $A \in \mathbf{H}(\text{tp})(\Gamma, h_\Gamma)$ to the set of sections of the projection $p_A : (\Gamma.A, h_{\Gamma.A}) \rightarrow (\Gamma, h_\Gamma)$ in \mathbf{H}_\diamond .

Construction 44 (*The hydration model*). With the definitions that we have given, the projection functor $\psi_\diamond : \mathbf{H}_\diamond \rightarrow \mathbf{A}_\diamond$ easily extends to a morphism of models, exhibiting \mathbf{H} as a atomic substitution model over \mathbf{A} .

Construction 45 (*The hydration map*). As we may compose $\psi : \mathbf{H} \rightarrow \mathbf{A}$ with the structure map $\alpha : \mathbf{A} \rightarrow \mathbf{I}$, we can view \mathbf{H} as a atomic substitution model over \mathbf{I} as well, and thus by the universal property of \mathbf{A} we have a universal section $\mathbf{A} \rightarrow \mathbf{H}$ to $\psi : \mathbf{H} \rightarrow \mathbf{A}$. Unraveling the section $\mathbf{A} \rightarrow \mathbf{H}$, we obtain precisely a natural assignment of hydration map component $h_\Gamma : \langle \Gamma \rangle \rightarrow \llbracket \Gamma \rrbracket$ from which we may assemble a single hydration map $\nearrow : \langle - \rangle \rightarrow \llbracket - \rrbracket$ in $\mathbf{Cat} \downarrow \mathbf{Sh} \mathbf{I}$.

6.8.4 Normalization and decidability

We can now show how to compute the normal form of a type $\Gamma \vdash A$ type, which we regard as an arrow $A : \alpha_\diamond \Gamma \rightarrow \text{tp}$ in \mathbb{T} . Then we may apply the functorial action of the evaluation functor $I_{\mathbf{Sh} \mathbf{G}} : \mathbb{T} \rightarrow \mathbf{Sh} \mathbf{G}$ to obtain $\llbracket A \rrbracket : \llbracket \Gamma \rrbracket \rightarrow \text{tp}^*$ and then postcompose with the projection of normal forms to obtain $\llbracket A \rrbracket.\text{code} : \llbracket \Gamma \rrbracket \rightarrow \text{nftp}$. Unraveling the meaning of such a map in the gluing category $\mathbf{Sh} \mathbf{G}$, we see that this amounts not to a normal form for A but instead to an assignment of normal forms of A to computability witnesses for all the variables in the context Γ . It is precisely

this gap that hydration fills:

$$\text{norm}_{\text{tp}}(\Gamma \vdash A \text{ type}) = \langle \Gamma \rangle \xrightarrow{\nearrow_{\Gamma}} \llbracket \Gamma \rrbracket \xrightarrow{\llbracket A \rrbracket} \text{tp}^* \xrightarrow{-.\text{code}} \text{nftp}$$

The entire map above restricts within object-space, by construction, to the original type A , or (to be more precise) its image in $\mathbf{Sh} \mathbf{G}$ under $j_* : \mathbf{Sh} \mathbf{I} \rightarrow \mathbf{Sh} \mathbf{G}$. The meta-space component of such a morphism $\langle \Gamma \rangle \rightarrow \text{nftp}$ is precisely a normal form for A .

Theorem 46. *Normalization is sound and complete:*

- (1) Soundness. *If $\text{norm}_{\text{tp}}(\Gamma \vdash A \text{ type}) = \text{norm}_{\text{tp}}(\Gamma \vdash B \text{ type})$, then $\Gamma \vdash A = B$.*
- (2) Completeness. *If $\Gamma \vdash A = B$, then $\text{norm}_{\text{tp}}(\Gamma \vdash A \text{ type}) = \text{norm}_{\text{tp}}(\Gamma \vdash B \text{ type})$.*

Proof. Completeness holds by definition, as the normalization function is defined on the denizens of the syntactic CwR rather than on raw terms. Soundness follows from the fact that $\text{norm}_{\text{tp}}(\Gamma \vdash A \text{ type})$ restricts in object-space to A itself. \square

In the same way, we can construct a normalization function for terms and prove that it is sound and complete (though we do not do so here). Of course, deciding equality for terms is practically important only insofar as it arises in the context of deciding equality for the types that mention them.

Definition 47. *An object $X \in \mathbf{Sh} \mathbf{G}$ has levelwise decidable equality when for each $\Gamma \in \mathbf{A}_{\diamond}$, the set $(i^*X)\Gamma$ has decidable equality where $i : A \hookrightarrow \mathbf{G}$ is as in Definition 19.*

Theorem 48. *Viewed as objects of $\mathbf{Sh} \mathbf{G}$, the following have levelwise decidable equality:*

$$\text{nftp} \quad (A : \text{tp}) \times \text{nf } A \quad (A : \text{tp}) \times (\phi : \Omega_{\text{dec}}) \times \text{ne}_{\phi} A$$

From this, we obtain our main results concerning $\mathbf{TT}_{\mathbb{P}}$:

Corollary 49. *Definitional equality in $\mathbf{TT}_{\mathbb{P}}$ is decidable.*

6.8.5 Stronger normalization results

A few stronger results can be proved using routine extensions of the methods on display here.

- (1) The external normalization function is surjective, which implies that normalization is idempotent. The main practical impact of normalization being surjective is to prove that the normalization function is effectively computable, as in Sterling (2021) and Sterling and Angiuli (2021); this step is redundant in the setting of the present paper, which has been carried out constructively in order to ensure an implicit form of effective computability.
- (2) Type constructors are injective in the sense that $\Gamma \vdash A \rightarrow B = A' \rightarrow B'$ implies $\Gamma \vdash A = A'$ and $\Gamma \vdash B = B'$, etc. Injectivity of type constructors is the main ingredient to establishing determinacy of the standard bidirectional elaboration algorithm. Injectivity is not strictly needed for a type checker written on fully annotated terms, but practical systems involve the elaboration of less-annotated terms to fully annotated terms; this process relies on injectivity.
- (3) Normalization can be internalized into $\mathbf{Sh} \mathbf{G}$ as an inverse $\text{tp} \rightarrow \text{nftp}$ to the projection $\text{nftp} \rightarrow \text{tp}$, as in Sterling (2021, 2025) and Sterling and Angiuli (2021). This implies, for example, that the normalization function is invariant under variable renaming (or, more generally, atomic substitutions).

We do not detail these results here, but instead remark that detailed proofs for similar theories can be found in the cited literature.

7. Related Work

Proof assistants already have support for various means of controlling the unfolding of definitions; we classify these as either library- or language-level.

7.1 Library-level features

Various library-level idioms for abstract definitions are used in practice such as SSReflect’s `lock` idiom. While such approaches are flexible and compatible with existing proof assistants, they are often cumbersome in practice. For instance, `lock` relies on various tactics with subtle behavior, which makes it difficult to use locking idioms in pure Gallina code.

7.2 Language-level features

Many proof assistants include a feature like Agda’s `abstract` blocks which marks a definition as completely opaque to the remainder of the development. In Remark 1, we explained how to recover Agda’s abstract definitions using controlled unfolding. Moreover, as controlled unfolding does not require a user to decide up front whether a definition can be unfolded, it gives a more realistic and flexible discipline for abstraction in a proof assistant. In practice, however, `abstract` is often used for performance reasons instead of merely for controlling abstraction; unfolding large or complex definitions can significantly slow down type checking and unification. While we have not discussed performance considerations for controlled unfolding, the same optimizations apply to our mechanism for definitions that are never unfolded. In total, controlled unfolding strictly generalizes `abstract` blocks.

Recently, Kovács (2023, 2024) has proposed a glued evaluation technique to both improve the pretty-printing of goals and more efficiently handle unfolding during conversion testing. Roughly, the proof assistant’s kernel may choose to unfold any definition but avoids doing so whenever possible for efficiency and strives to never show unfolded goals to the user. Both glued evaluation and controlled unfolding relate to the unfolding of definitions, and they are largely orthogonal and complementary. In particular, glued evaluation does not require user intervention, unlike controlled unfolding, but it does not actually preclude any unfolding from taking place. Thus, glued evaluation does not impact the well-formedness of a program and can be used as a “drop-in” technique for improving performance and usability. However, for the same reasons, glued evaluation cannot be used to enforce modularity and independence in the same way as controlled unfolding. Ideally, a proof assistant would support both glued evaluation and controlled unfolding: the more advanced evaluation algorithm would improve baseline performance and controlled unfolding would facilitate users enforcing stronger abstraction boundaries within their programs and assisting the kernel by manually designating certain definitions as opaque.

Program verifiers such as VeriFast and Chalice include similar unfolding mechanisms to cope specifically with recursive definitions (Jacobs et al. 2015; Summers and Drossopoulou 2013). Like our mechanism, these features allow users fine-grained control over how definitions are unfolded. However, these verifiers work only within simply typed theories and thus avoid the substantial complexity of dependency. Moreover, these mechanisms manage a different problem than controlled unfolding; they allow a user to unfold recursive definitions step-by-step, while controlled unfolding is used to control when each definition can be fully inlined.

7.3 Translucent ascription in module systems

Thus far we have focused on proof assistants, but similar considerations arise for ML-style module systems (Dreyer et al. 2003; Harper and Stone 2000; Milner et al. 1997; Sterling and Harper 2021). The default opacity for definitions in module systems is the same as in controlled unfolding and opposite to proof assistants: types are abstract unless marked otherwise. The treatment of translucent type declarations in module systems (Harper and Stone 2000) relies on singleton kinds (Aspinall 1995; Stone and Harper 2006), which are the special case of extension types whose boundary proposition is \top . Generalizing from compiletime kinds to mixed compiletime–runtime module signatures, Sterling and Harper have pointed out that transparent ascriptions are best handled by an extension type whose boundary proposition represents the compiletime phase itself (Sterling and Harper 2021). Thus, the translucency of compiletime module components can be seen as a particular controlled unfolding policy in the sense of this paper.

7.4 Controlled unfolding in Agda

Inspired by our implementation of controlled unfolding in `cooltt`, Amélia Liao and Jesper Cockx have implemented a version of this mechanism called `opaque` within Agda 2.6.4 (Liao and Cockx 2022). However, rather than using extension types, their Agda implementation simulates the necessary behaviors by instrumenting conversion checking – a workaround made possible by the very restricted ways in which our elaboration procedure relies on extension types. This demonstrates that controlled unfolding can be adapted to proof assistants like Coq whose core calculi do not presently support extension types.

At the time of writing, Agda’s `opaque` declarations are new enough that only two major Agda libraries, the 1Lab (The 1Lab Development Team 2022) and the Cubical Agda library (The Agda Community 2023), use them extensively; the Agda standard library may adopt `opaque` declarations in a future major revision (The Agda Development Team 2023). As of publication, 65 modules in the 1Lab use `opaque` and 19 use `unfolding`, in addition to over 100 using `abstract` blocks; in the Cubical Agda library, 27 modules use `opaque`, 6 use `unfolding`, and 35 use `abstract`.

8. Conclusions and Future Work

We have proposed controlled unfolding, a new mechanism for interpolating between transparent and opaque definitions in proof assistants. We have demonstrated its practical applicability by extending `cooltt` with controlled unfolding; we have also proved its soundness through an elaboration algorithm to a core calculus whose normalization we establish using a constructive STC argument.

In the future, we hope to see controlled unfolding integrated into more proof assistants and to further explore its applications for large-scale organization of mechanized mathematics. As mentioned above, some our mechanism has implemented in Agda, but features such as local unfolds are still absent. Furthermore, in the context of our `cooltt` implementation, we have also already begun to experiment with potential extensions, including one that allows a subterm to be declared locally abstract and then unfolded later on as needed – a more flexible alternative to Coq’s `abstract t` tactical. As we mentioned in Remark 1, we also are interested in facilities to limit the scope in which it is possible to unfold a definition.

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Notes

- 1 Indeed, the Agda standard library (The Agda Development Team 2022) currently uses `abstract` only once.
- 2 <https://github.com/coq/coq/blob/V8.16.0/theories/ssr/ssreflect.v/#L388>
- 3 By *constructive*, we mean something that can be carried out in an elementary topos with a natural numbers object.
- 4 In his doctoral dissertation, Sterling referred to “object-space” and “meta-space” as *syntactic* and *semantic*, respectively (Sterling 2021). However, there are compelling reasons to consider object-space more semantic than meta-space (in which various admissibilities hold that will not be preserved by homomorphisms of models), so we have changed terminology to avoid confusion.
- 5 It would be possible to choose a more restrictive class of representable maps for **Sh G**, but there is no reason to do so.

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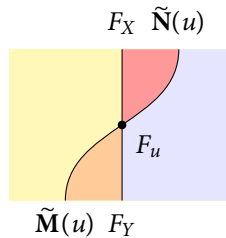
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A. The (2,1)-category of models

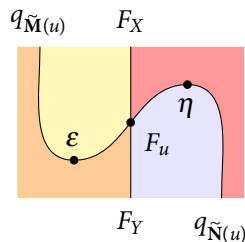
Uemura (2021) has observed that a model $(\mathbf{M}_\diamond, \mathbf{M})$ in the sense of Definition 11 can be packaged into a single functor $\tilde{\mathbf{M}}: \mathbb{T}^\triangleright \rightarrow \mathbf{Cat}$, in which $\mathbb{T}^\triangleright$ freely extends \mathbb{T} by a new terminal object \diamond and \mathbf{Cat} is the 2-category of categories. From this perspective, a sort $X \in \mathbb{T}$ is taken to the total category $\tilde{\mathbf{M}}(X) = \int_{\mathbf{M}_\diamond} \mathbf{M}(X)$ of a discrete fibration over $\tilde{\mathbf{M}}(\diamond) = \mathbf{M}_\diamond$. Here, we are using the equivalence between $\mathbf{DFib}_{\mathcal{C}} \simeq \mathbf{Pr} \mathcal{C}$. The preservation of representable maps is then rendered here as the requirement that for representable $u: X \rightarrow Y$, each functor $\tilde{\mathbf{M}}(u): \tilde{\mathbf{M}}(X) \rightarrow \tilde{\mathbf{M}}(Y)$ shall have a right adjoint $\tilde{\mathbf{M}}(u) \dashv q_{\tilde{\mathbf{M}}(u)}$ taking an element of $\tilde{\mathbf{M}}(Y)$ to the *generic* element of $\tilde{\mathbf{M}}(X)$ in the extended context.

Example 50. For the representable map $\pi: \mathbf{tm} \rightarrow \mathbf{tp}$, the functorial action $\tilde{\mathbf{M}}(\pi): \tilde{\mathbf{M}}(\mathbf{tm}) \rightarrow \tilde{\mathbf{M}}(\mathbf{tp})$ takes a term $\Gamma \vdash a: A$ to the type $\Gamma \vdash A$; the right adjoint $q_{\tilde{\mathbf{M}}(\pi)}: \tilde{\mathbf{M}}(\mathbf{tp}) \rightarrow \tilde{\mathbf{M}}(\mathbf{tm})$ sends a type $\Gamma \vdash A$ to the variable $\Gamma, a: A \vdash a: A$.

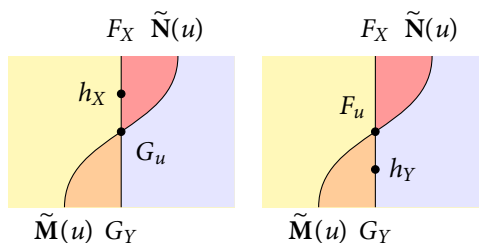
Definition 51. Given two models \mathbf{M}, \mathbf{N} of \mathbb{T} , a morphism of models from \mathbf{M} to \mathbf{N} is given by a natural transformation $F: \tilde{\mathbf{M}} \rightarrow \tilde{\mathbf{N}} \in [\mathbb{T}^\triangleright, \mathbf{Cat}]$ such that for each representable map $u: X \rightarrow Y$ in \mathbb{T} , the corresponding naturality datum $F_u: \tilde{\mathbf{N}}(u) \circ F_X = F_Y \circ \tilde{\mathbf{M}}(u)$ depicted below



satisfies the Beck–Chevalley condition in the sense the following 2-cell, obtained by conjugating with units and counits, denotes an invertible natural transformation $F_X \circ q_{\tilde{\mathbf{M}}(u)} \rightarrow q_{\tilde{\mathbf{N}}(u)} \circ F_Y$:



Definition 52. Let \mathbf{M}, \mathbf{N} be two models of \mathbb{T} , and let $F, G: \mathbf{M} \rightarrow \mathbf{N}$ be two morphisms of models. An isomorphism h from F to G is defined to be an invertible modification between the underlying natural transformations F, G . This amounts to choosing for each $X \in \mathbb{T}^\triangleright$ a natural isomorphism $h_X: F_X \rightarrow G_X$ in $[\tilde{\mathbf{M}}(X), \tilde{\mathbf{N}}(X)]$, subject to the coherence condition that for each $u: X \rightarrow Y$ in $\mathbb{T}^\triangleright$ the following two wiring diagrams are equal:



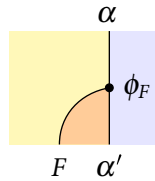
Remark 53. Because each of the induced maps $\pi_{\mathbf{M}(X)} : \tilde{\mathbf{M}}(X) \rightarrow \mathbf{M}_\diamond$ and $\pi_{\mathbf{N}(X)} : \tilde{\mathbf{N}}(X) \rightarrow \mathbf{N}_\diamond$ into the cone point are discrete fibrations, it suffices to check the modification condition of h on only the cone maps $X \rightarrow \diamond$: as any discrete fibration is a faithful functor, it moreover follows that $h_\diamond : F_\diamond \rightarrow G_\diamond$ uniquely determines all the other h_X if they exist. Unfolding further, given $x \in \tilde{\mathbf{M}}(X)$ we are only requiring that $(h_\diamond)^*_{\pi_{\mathbf{M}(X)}}(G_X x) = F_X x$ in the sense depicted below in the discrete fibration $\mathbf{N}(X)$ over \mathbf{N}_\diamond :

$$\begin{array}{ccc}
 F_X x & \xrightarrow{\exists! h_X x} & G_X x \\
 \downarrow & \lrcorner & \downarrow \\
 F_\diamond(\pi_{\mathbf{M}(X)} x) & \xrightarrow{(h_\diamond)_{\pi_{\mathbf{M}(X)}}} & G_\diamond(\pi_{\mathbf{M}(X)} x)
 \end{array}$$

Thus, we have a (2,1)-category of models $\mathbf{Mod} \mathbb{T}$ for any category with representable maps \mathbb{T} .

B. The (2,1)-category of atomic substitution models

Definition 54. Given atomic substitution models $\alpha : \mathbf{A} \rightarrow \mathbf{I}$ and $\alpha' : \mathbf{A}' \rightarrow \mathbf{I}$, a morphism from (\mathbf{A}, α) to (\mathbf{A}', α') is given by a morphism $F : \mathbf{A} \rightarrow \mathbf{A}' \in \mathbf{Mod} \mathbb{T}_0$ together with an isomorphism $\phi_F : \alpha \rightarrow \alpha' \circ F$ in $[\mathbf{A}, \mathbf{I}]$ as depicted below:



Definition 55. Given two morphisms $F, G : (\mathbf{A}, \alpha) \rightarrow (\mathbf{A}', \alpha')$, an isomorphism from F to G is given by an isomorphism $h : F \rightarrow G \in [\mathbf{A}, \mathbf{A}']$ such that the following wiring diagrams denote equal isomorphisms $\alpha \rightarrow \alpha' \circ G$:

