

Type-Theoretic Signatures for Algebraic Theories and Inductive Types

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CHAPTER 1

Introduction

1.1 Specification and Semantics for Inductive Types

1.2 Overview of the Thesis and Contributions

1.3 Notation and Conventions

1.3.1 Metatheory

1.3.2 Universes

CHAPTER 2

Simple Inductive Signatures

In this chapter, we take a look at a very simple notion of inductive signature. The motivation for doing so is to present the basic ideas of this thesis in the easiest possible setting, with explicit definitions. The later chapters are greatly generalized and expanded compared to the current one, and are not feasible (and probably not that useful) to present in full formal detail. We also include a complete Agda formalization of the contents of this chapter, in less than 200 lines.

potentially in intro

The mantra throughout this dissertation is the following: inductive types are specified by typing contexts in certain *theories of signatures*. For each class of inductive types, there is a corresponding theory of signatures, which is viewed as a proper type theory and comes equipped with an algebraic model theory. *Semantics* of signatures is given by interpreting them in certain models of the theory of signatures. Semantics should at least provide a notion of induction principle for each signature; in this chapter we provide a bit more than that, and substantially more in Chapters 4 and 5.

2.1 Theory of Signatures

Generally, more expressive theories of signatures can describe a larger classes of inductive types. As we are aiming at minimalism right now, the current theory of signatures is as follows:

Definition 1. The **theory of signatures**, or ToS for short, is a simple type theory equipped with the following features:

- An empty base type ι .
- A *first-order function type* $\iota \rightarrow -$; this is a function whose domain is fixed to be ι . Moreover, first-order functions only have neutral terms: there is application, but no λ -abstraction.

We can specify the full syntax using the following Agda-like inductive definitions.

$$\begin{array}{ll}
\text{Ty} & : \text{Set} \\
\iota & : \text{Ty} \\
\iota \rightarrow - & : \text{Ty} \rightarrow \text{Ty} \\
\\
\text{Var} & : \text{Con} \rightarrow \text{Ty} \rightarrow \text{Set} \\
\text{vz} & : \text{Var } (\Gamma \triangleright A) A \\
\text{vs} & : \text{Var } \Gamma A \rightarrow \text{Var } (\Gamma \triangleright B) A \\
\\
\text{Con} & : \text{Set} \\
\bullet & : \text{Con} \\
- \triangleright - & : \text{Con} \rightarrow \text{Ty} \rightarrow \text{Con} \\
\\
\text{Tm} & : \text{Con} \rightarrow \text{Ty} \rightarrow \text{Set} \\
\text{var} & : \text{Var } \Gamma A \rightarrow \text{Tm } \Gamma A \\
\text{app} & : \text{Tm } \Gamma (\iota \rightarrow A) \rightarrow \text{Tm } \Gamma \iota \rightarrow \text{Tm } \Gamma A
\end{array}$$

Here, Con contexts are lists of types, and Var specifies well-typed De Bruijn indices, where vz represents the zero index, and vs takes the successor of an index.

Notation 1. We use capital Greek letters starting from Γ to refer to contexts, A, B, C to refer to types, and t, u, v to refer to terms. In examples, we may use a nameful notation instead of De Bruijn indices. For example, we may write $x : \text{Tm } (\bullet \triangleright (x : \iota) \triangleright (y : \iota)) \iota$ instead of $\text{var } (\text{vs vz}) : \text{Tm } (\bullet \triangleright \iota \triangleright \iota) \iota$. Additionally, we may write $t u$ instead of $\text{app } t u$ for t and u terms.

Definition 2. **Parallel substitutions** map variables to terms.

$$\begin{array}{l}
\text{Sub} : \text{Con} \rightarrow \text{Con} \rightarrow \text{Set} \\
\text{Sub } \Gamma \Delta \equiv \{A\} \rightarrow \text{Var } \Delta A \rightarrow \text{Tm } \Gamma A
\end{array}$$

We use σ and δ to refer to substitutions. We also recursively define the action of substitution on terms:

$$\begin{array}{l}
-[-] : \text{Tm } \Delta A \rightarrow \text{Sub } \Gamma \Delta \rightarrow \text{Tm } \Gamma A \\
(\text{var } x) \ [-\sigma] \equiv \sigma x \\
(\text{app } t u) [\sigma] \equiv \text{app } (t[\sigma]) (u[\sigma])
\end{array}$$

The identity substitution id is defined simply as var . It is easy to see that $t[\text{id}] = t$ for all t . Substitution composition is as follows.

$$\begin{aligned} - \circ - &: \text{Sub } \Delta \Xi \rightarrow \text{Sub } \Gamma \Delta \rightarrow \text{Sub } \Gamma \Xi \\ (\sigma \circ \delta) x &:= (\sigma x)[\delta] \end{aligned}$$

Example 1. We may write signatures for natural numbers and binary trees respectively as follows.

$$\begin{aligned} \text{NatSig} &:= \bullet \triangleright (\text{zero} : \iota) \triangleright (\text{suc} : \iota \rightarrow \iota) \\ \text{TreeSig} &:= \bullet \triangleright (\text{leaf} : \iota) \triangleright (\text{node} : \iota \rightarrow \iota \rightarrow \iota) \end{aligned}$$

In short, the current ToS allows inductive types which are

- *Single-sorted*: this means that we have a single type constructor, corresponding to ι .
- *Closed*: signatures cannot refer to any externally existing type. For example, we cannot write a signature for “lists of natural number” in a direct fashion, since there is no way to refer to the type of natural numbers.
- *Finitary*: inductive types corresponding to signatures are always finitely branching trees. Being closed implies being finitary, since an infinitely branching node would require some external type to index subtrees with. For example, $\text{node} : (\mathbb{N} \rightarrow \iota) \rightarrow \iota$ would specify an infinite branching (if such type was allowed in ToS).

Remark. We omit λ -expressions from ToS for the sake of simplicity: this causes terms to be always in normal form (neutral, to be precise), and thus we can skip talking about conversion rules. Later, starting from Chapter 4 we include proper $\beta\eta$ -rules in signature theories.

2.2 Semantics

For each signature, we need to know what it means for a type theory to support the corresponding inductive type. For this, we need at least a notion of *algebras*, which can be viewed as a bundle of all type and value constructors, and what it means for an algebra to support an *induction principle*. Additionally, we may

want to know what it means to support a *recursion principle*, which can be viewed as a non-dependent variant of induction. In the following, we define these notions by induction on ToS syntax.

Remark. We use “algebra” and “model” synonymously throughout this thesis.

2.2.1 Algebras

First, we calculate types of algebras. This is simply a standard interpretation into the **Set** universe. We define the following operations by induction; the $-^A$ name is overloaded for **Con**, **Ty** and **Tm**.

$$-^A : \mathbf{Ty} \rightarrow \mathbf{Set} \rightarrow \mathbf{Set}$$

$$\iota^A \quad X : \equiv X$$

$$(\iota \rightarrow A)^A X : \equiv X \rightarrow A^A X$$

$$-^A : \mathbf{Con} \rightarrow \mathbf{Set} \rightarrow \mathbf{Set}$$

$$\Gamma^A X : \equiv \{A : \mathbf{Ty}\} \rightarrow \mathbf{Var} \Gamma A \rightarrow A^A X$$

$$-^A : \mathbf{Tm} \Gamma A \rightarrow \{X : \mathbf{Set}\} \rightarrow \Gamma^A X \rightarrow A^A X$$

$$(\mathbf{var} \ x)^A \quad \gamma : \equiv \gamma \ x$$

$$(\mathbf{app} \ t \ u)^A \gamma : \equiv t^A \gamma (u^A \gamma)$$

$$-^A : \mathbf{Sub} \Gamma \Delta \rightarrow \{X : \mathbf{Set}\} \rightarrow \Gamma^A X \rightarrow \Delta^A X$$

$$\sigma^A \gamma \ x : \equiv (\sigma \ x)^A \gamma$$

Here, types and contexts depend on some $X : \mathbf{Set}$, which serves as the interpretation of ι . We define Γ^A as a product: for each variable in the context, we get a semantic type. This trick, along with the definition of **Sub**, makes formalization a bit more compact. Terms and substitutions are interpreted as natural maps. Substitutions are interpreted by pointwise interpreting the contained terms.

Notation 2. We may write values of Γ^A using notation for Σ -types. For example, we may write $(\mathbf{zero} : X) \times (\mathbf{suc} : X \rightarrow X)$ for the result of computing $\mathbf{NatSig}^A X$.

Definition 3. We define **algebras** as follows.

$$\begin{aligned} \text{Alg} &: \text{Con} \rightarrow \text{Set}_1 \\ \text{Alg } \Gamma &::= (X : \text{Set}) \times \Gamma^A X \end{aligned}$$

Example 2. Alg NatSig is computed to $(X : \text{Set}) \times (\text{zero} : X) \times (\text{suc} : X \rightarrow X)$.

2.2.2 Morphisms

Now, we compute notions of morphisms of algebras. In this case, morphisms are functions between underlying sets which preserve all specified structure. The interpretation for calculating morphisms is a *logical relation interpretation* [HRR14] over the $-^A$ interpretation. The key part is the interpretation of types:

$$\begin{aligned} -^M &: (A : \text{Ty})\{X_0 X_1 : \text{Set}\}(X^M : X_0 \rightarrow X_1) \rightarrow A^A X_0 \rightarrow A^A X_1 \rightarrow \text{Set} \\ \iota^M & \quad X^M \alpha_0 \alpha_1 ::= X^M \alpha_0 = \alpha_1 \\ (\iota \rightarrow A)^M & X^M \alpha_0 \alpha_1 ::= (x : X_0) \rightarrow A^M X^M (\alpha_0 x) (\alpha_1 (X^M x)) \end{aligned}$$

We again assume an interpretation for the base type ι , as X_0 , X_1 and $X^M : X_0 \rightarrow X_1$. X^M is function between underlying sets of algebras, and A^M computes what it means that X^M preserves an operation with type A . At the base type, preservation is simply equality. At the first-order function type, preservation is a quantified statement over X_0 . We define morphisms for **Con** pointwise:

$$\begin{aligned} -^M &: (\Gamma : \text{Con})\{X_0 X_1 : \text{Set}\} \rightarrow (X_0 \rightarrow X_1) \rightarrow \Gamma^A X_0 \rightarrow \Gamma^A X_1 \rightarrow \text{Set} \\ \Gamma^M X^M \gamma_0 \gamma_1 &::= \{A : \text{Ty}\}(x : \text{Var } \Gamma A) \rightarrow A^M X^M (\gamma_0 x) (\gamma_1 x) \end{aligned}$$

For terms and substitutions, we get preservation statements, which are sometimes called *fundamental lemmas* in discussions of logical relations [HRR14].

$$\begin{aligned} -^M &: (t : \text{Tm } \Gamma A) \rightarrow \Gamma^M X^M \gamma_0 \gamma_1 \rightarrow A^M X^M (t^A \gamma_0) (t^A \gamma_1) \\ (\text{var } x)^M & \quad \gamma^M ::= \gamma^M x \\ (\text{app } t u)^M & \gamma^M ::= t^M \gamma^M (u^A \gamma_0) \end{aligned}$$

$$\begin{aligned} -^M &: (\sigma : \text{Sub } \Gamma \Delta) \rightarrow \Gamma^M X^M \gamma_0 \gamma_1 \rightarrow \Delta^M X^M (\sigma^A \gamma_0) (\sigma^A \gamma_1) \\ \sigma^M \gamma^M x &::= (\sigma x)^M \gamma^M \end{aligned}$$

The definition of $(\text{app } t u)^M$ is well-typed by the induction hypothesis $u^M \gamma^M : X^M (u^A \gamma_0) = u^A \gamma_1$.

Definition 4. To get notions of **algebra morphisms**, we again pack up Γ^M with the interpretation of ι .

$$\begin{aligned} \text{Mor} : \{\Gamma : \text{Con}\} &\rightarrow \text{Alg } \Gamma \rightarrow \text{Alg } \Gamma \rightarrow \text{Set} \\ \text{Mor } \{\Gamma\} (X_0, \gamma_0) (X_1, \gamma_1) &\equiv (X^M : X_0 \rightarrow X_1) \times \Gamma^M X^M \gamma_0 \gamma_1 \end{aligned}$$

Example 3. We have the following computation:

$$\begin{aligned} \text{Mor } \{\text{NatSig}\} (X_0, \text{zero}_0, \text{suc}_0) (X_0, \text{zero}_1, \text{suc}_1) &\equiv \\ (X^M : X_0 \rightarrow X_1) & \\ \times (X^M \text{zero}_0 = \text{zero}_1) & \\ \times ((x : X_0) \rightarrow X^M (\text{suc}_0 x) = \text{suc}_1 (X^M x)) & \end{aligned}$$

Definition 5. We state **initiality** as a predicate on algebras:

$$\begin{aligned} \text{Initial} : \{\Gamma : \text{Con}\} &\rightarrow \text{Alg } \Gamma \rightarrow \text{Set} \\ \text{Initial } \{\Gamma\} \gamma &\equiv (\gamma' : \text{Alg } \Gamma) \rightarrow \text{isContr } (\text{Mor } \gamma \gamma') \end{aligned}$$

Here `isContr` refers to unique existence [Uni13, Section 3.11]. If we drop `isContr` from the definition, we get the notion of weak initiality, which corresponds to the recursion principle for Γ . Although we call this predicate **Initial**, in this chapter we do not yet show that algebras form a category. We provide the extended semantics in Chapter 4. The computed algebras and morphism there remain the same as in the current chapter.

Morphisms vs. logical relations. The $-^M$ interpretation can be viewed as a special case of logical relations over the $-^A$ model: every morphism is a *functional* logical relation, where the chosen relation between the underlying sets happens to be a function. Consider now a more general relational interpretation for types:

$$\begin{aligned} -^R : (A : \text{Ty}) \{X_0 X_1 : \text{Set}\} (X^R : X_0 \rightarrow X_1 \rightarrow \text{Set}) &\rightarrow A^A X_0 \rightarrow A^A X_1 \rightarrow \text{Set} \\ \iota^R \quad X^R \alpha_0 \alpha_1 &\equiv X^R \alpha_0 \alpha_1 \\ (\iota \rightarrow A)^R X^R \alpha_0 \alpha_1 &\equiv (x_0 : X_0)(x_1 : X_1) \rightarrow X^R x_0 x_1 \rightarrow A^R X^R (\alpha_0 x_0) (\alpha_1 x_1) \end{aligned}$$

Here, functions are related if they map related inputs to related outputs. If we know that $X^M \alpha_0 \alpha_1 \equiv (f \alpha_0 = \alpha_1)$ for some f function, we get

$$(x_0 : X_0)(x_1 : X_1) \rightarrow f x_0 = x_1 \rightarrow A^R X^R (\alpha_0 x_0) (\alpha_1 x_1)$$

Now, we can simply substitute along the input equality proof in the above type, to get the previous definition for $(\iota \rightarrow A)^M$:

$$(x_0 : X_0) \rightarrow A^R X^R (\alpha_0 x_0) (\alpha_1 (f x_0))$$

This substitution along the equation is called “singleton contraction” in the jargon of homotopy type theory [Uni13]. The ability to perform contraction here is at the heart of the *strict positivity restriction* for inductive signatures. Strict positivity in our setting corresponds to only having first-order function types in signatures. If we allowed function domains to be arbitrary types, in the definition of $(A \rightarrow B)^M$ we would only have a black-box $A^M X^M : A^A X_0 \rightarrow A^A X_1 \rightarrow \mathbf{Set}$ relation, which is not known to be given as an equality.

In Chapter 4 we expand on this. As a preliminary summary: although higher-order functions have relational interpretation, such relations do not generally compose. What we eventually aim to have is a *category* of algebras and algebra morphisms, where morphisms do compose. We need a *directed* model of the theory of signatures, where every signature becomes a category of algebras. The way to achieve this, is to prohibit higher-order functions, thereby avoiding the polarity issues that prevent a directed interpretation for general function types.

2.2.3 Displayed Algebras

At this point we do not yet have specification for induction principles. We use the term *displayed algebra* to refer to “dependent” algebras, where every displayed algebra component lies over corresponding components in the base algebra. For the purpose of specifying induction, displayed algebras can be viewed as bundles of induction motives and methods.

Displayed algebras over some $\gamma : \mathbf{Alg} \Gamma$ are equivalent to slices over γ in the category of Γ -algebras; we show this in Chapter 4. A slice $f : \mathbf{Sub} \Gamma \gamma' \gamma$ maps elements of γ ’s underlying set to elements in the base algebra. Why do we need displayed algebras, then? The main reason is that if we are to eventually implement inductive types in a dependently typed language, we need to compute induction principles exactly, not merely up to isomorphisms.

For more illustration of using some displayed algebras in a type-theoretic setting, see [AL19]. We adapt the term “displayed algebra” from *ibid.* as a generalization of displayed categories (and functors, natural transformations) to other algebraic structures.

The displayed algebra interpretation is a *logical predicate* interpretation, defined as follows.

$$\begin{aligned} -^D &: (A : \mathbf{Ty})\{X\} \rightarrow (X \rightarrow \mathbf{Set}) \rightarrow A^A X \rightarrow \mathbf{Set} \\ \iota^D \quad X^D \alpha &\equiv X^D \alpha \\ (\iota \rightarrow A)^D X^D \alpha &\equiv (x : X)(x^D : X^D x) \rightarrow A^D X^D (\alpha x) \end{aligned}$$

$$\begin{aligned} -^D &: (\Gamma : \mathbf{Con})\{X\} \rightarrow (X \rightarrow \mathbf{Set}) \rightarrow \Gamma^A X \rightarrow \mathbf{Set} \\ \Gamma^D X^D \gamma &\equiv \{A : \mathbf{Ty}\}(x : \mathbf{Var} \Gamma A) \rightarrow A^D X^D (\gamma x) \end{aligned}$$

$$\begin{aligned} -^D &: (t : \mathbf{Tm} \Gamma A) \rightarrow \Gamma^D X^D \gamma \rightarrow A^D X^D (t^A \gamma) \\ (\mathbf{var} x)^D \gamma^D &\equiv \gamma^D x \\ (\mathbf{app} t u)^D \gamma^D &\equiv t^D \gamma^D (u^A \gamma) (u^D \gamma^D) \end{aligned}$$

$$\begin{aligned} -^D &: (\sigma : \mathbf{Sub} \Gamma \Delta) \rightarrow \Gamma^D X^D \gamma \rightarrow \Delta^D X^D (\sigma^A \gamma) \\ \sigma^D \gamma^D x &\equiv (\sigma x)^D \gamma^D \end{aligned}$$

Analogously to before, everything depends on a predicate interpretation $X^D : X \rightarrow \mathbf{Set}$ for ι . For types, a predicate holds for a function if the function preserves predicates. The interpretation of terms is again a fundamental lemma, and we again have pointwise definitions for contexts and substitutions.

Definition 6 (displayed algebras).

$$\begin{aligned} \mathbf{DispAlg} &: \{\Gamma : \mathbf{Con}\} \rightarrow \mathbf{Alg} \Gamma \rightarrow \mathbf{Set}_1 \\ \mathbf{DispAlg} \{\Gamma\} (X, \gamma) &\equiv (X^D : X \rightarrow \mathbf{Set}) \times \Gamma^D X^D \gamma \end{aligned}$$

Example 4. We have the following computation.

$$\begin{aligned} \mathbf{DispAlg} \{\mathbf{NatSig}\} (X, \mathit{zero}, \mathit{suc}) &\equiv \\ & (X^D : X \rightarrow \mathbf{Set}) \\ & \times (\mathit{zero}^D : X^D \mathit{zero}) \\ & \times (\mathit{suc}^D : (n : X) \rightarrow X^D n \rightarrow X^D (\mathit{suc} n)) \end{aligned}$$

2.2.4 Sections

Sections of displayed algebras are “dependent” analogues of algebra morphisms, where the codomain is displayed over the domain.

$$\begin{aligned} -^S &: (A : \mathbf{Ty})\{X X^D\}(X^S : (x : X) \rightarrow X^D x) \rightarrow (\alpha : A^A X) \rightarrow A^D X^D \alpha \rightarrow \mathbf{Set} \\ \iota^S & \quad X^S \alpha \alpha^D \equiv X^S \alpha = \alpha^D \\ (\iota \rightarrow A)^S & X^S \alpha \alpha^D \equiv (x : X) \rightarrow A^S X^S (\alpha x) (\alpha^D (X^S x)) \end{aligned}$$

$$\begin{aligned} \mathbf{Con}^S &: (\Gamma : \mathbf{Con})\{X X^D\}(X^S : (x : X) \rightarrow X^D x) \rightarrow (\gamma : \Gamma^A X) \rightarrow \Gamma^D X^D \gamma \rightarrow \mathbf{Set} \\ \Gamma^S X^S \gamma_0 \gamma_1 &\equiv \{A : \mathbf{Ty}\}(x : \mathbf{Var} \Gamma A) \rightarrow A^S X^S (\gamma_0 x) (\gamma_1 x) \end{aligned}$$

$$\begin{aligned} -^S &: (t : \mathbf{Tm} \Gamma A) \rightarrow \Gamma^S X^S \gamma \gamma^D \rightarrow A^S X^S (t^A \gamma) (t^D \gamma^D) \\ (\mathbf{var} x)^S & \gamma^S \equiv \gamma^S x \\ (\mathbf{app} t u)^S & \gamma^S \equiv t^S \gamma^S (u^A \gamma) \end{aligned}$$

$$\begin{aligned} -^S &: (\sigma : \mathbf{Sub} \Gamma \Delta) \rightarrow \Gamma^S X^S \gamma \gamma^D \rightarrow \Delta^S X^S (\sigma^A \gamma) (\sigma^A \gamma^D) \\ \sigma^S \gamma^S x &= (\sigma x)^S \gamma^S \end{aligned}$$

Definition 7 (Displayed algebra sections (“sections” in short)).

$$\begin{aligned} \mathbf{Section} &: \{\Gamma : \mathbf{Con}\} \rightarrow (\gamma : \mathbf{Alg} \Gamma) \rightarrow \mathbf{DispAlg} \gamma \rightarrow \mathbf{Set} \\ \mathbf{Section} (X, \gamma) (X^D \gamma^D) &\equiv (X^S : (x : X) \rightarrow X^D x) \times \Gamma^S X^S \gamma \gamma^D \end{aligned}$$

Example 5. We have the following computation.

$$\begin{aligned} \mathbf{Section} \{\mathbf{NatSig}\} (X, \mathbf{zero}, \mathbf{suc}) (X^D, \mathbf{zero}^D, \mathbf{suc}^D) &\equiv \\ & (X^S : (x : X) \rightarrow X^D x) \\ & \times (\mathbf{zero}^S : X^S \mathbf{zero} = \mathbf{zero}^D) \\ & \times (\mathbf{suc}^S : (n : X) \rightarrow X^S (\mathbf{suc} n) = \mathbf{suc}^D n (X^S n)) \end{aligned}$$

Definition 8 (Induction). We define a predicate which holds if an algebra supports induction.

$$\begin{aligned} \mathbf{Inductive} &: \{\Gamma : \mathbf{Con}\} \rightarrow \mathbf{Alg} \Gamma \rightarrow \mathbf{Set}_1 \\ \mathbf{Inductive} \{\Gamma\} \gamma &\equiv (\gamma^D : \mathbf{DispAlg} \gamma) \rightarrow \mathbf{Section} \gamma \gamma^D \end{aligned}$$

We can observe that $\text{Inductive}\{\text{NatSig}\}(X, \text{zero}, \text{suc})$ computes exactly to the usual induction principle for natural numbers. The input DispAlg is a bundle of the induction motive and the methods, and the output Section contains the X^S eliminator function together with its β computation rules.

2.3 Term Algebras

In this section we show that if a type theory supports the inductive types comprising the theory of signatures, it also supports every inductive type which is described by the signatures.

Note that we specified Tm and Sub , but did not need either of them when specifying signatures, or when computing induction principles. That signatures do not depend on terms, is a property specific to simple signatures; this will not be the case in Chapter 4 when we move to more general signatures. However, terms and substitutions are already useful here in the construction of term algebras.

The idea is that terms in contexts comprise initial algebras. For example, $\text{Tm NatSig } \iota$ is the set of natural numbers (up to isomorphism). Informally, this is because the only way to construct terms is by applying the suc variable (given by var vz) finitely many times to the zero variable (given by var (vs vz)).

Definition 9 (Term algebras). Fix an $\Omega : \text{Con}$. We abbreviate $\text{Tm } \Omega \iota$ as T ; this will serve as the carrier set of the term algebra. We additionally define the following.

$$-^T : (A : \text{Ty}) \rightarrow \text{Tm } \Omega A \rightarrow A^A \text{T}$$

$$\iota^T \quad t \equiv t$$

$$(\iota \rightarrow A)^T t \equiv \lambda u. A^T (\text{app } t u)$$

$$-^T : (\Gamma : \text{Con}) \rightarrow \text{Sub } \Omega \Gamma \rightarrow \Gamma^A \text{T}$$

$$\Gamma^T \nu \{A\} x \equiv A^T (\nu x)$$

$$-^T : (t : \text{Tm } \Gamma A)(\nu : \text{Sub } \Omega \Gamma) \rightarrow A^T (t[\nu]) = t^A (\Gamma^T \nu)$$

$$(\text{var } x)^T \quad \nu \text{ holds by refl}$$

$$(\text{app } t u)^T \nu \text{ holds by } t^T \nu \text{ and } u^T \nu$$

$$\begin{aligned}
& -^T : (\sigma : \mathbf{Sub} \Gamma \Delta)(\nu : \mathbf{Sub} \Omega \Gamma)\{A\}(x : \mathbf{Var} \Delta A) \\
& \quad \rightarrow \Delta^T (\sigma \circ \nu) x = \sigma^A (\Gamma^T \nu) x \\
& \sigma^T \nu x \equiv (\sigma x)^T \nu
\end{aligned}$$

Now we can define the term algebra for Ω itself:

$$\begin{aligned}
& \mathbf{TmAlg}_\Omega : \mathbf{Alg} \Omega \\
& \mathbf{TmAlg}_\Omega \equiv \Omega^T \Omega \text{id}
\end{aligned}$$

In the interpretation for contexts, it is important that Ω is fixed, and we do induction on all Γ contexts such that there is a $\mathbf{Sub} \Omega \Gamma$. It would not work to try to compute term algebras by direct induction on contexts, because we need to refer to the same \mathbf{T} set in the interpretation of every type in a signature.

The interpretation of types embeds terms as A -algebras. At the base type ι , this embedding is simply the identity function, since $\iota^A \mathbf{T} \equiv \mathbf{T} \equiv \mathbf{Tm} \Omega \iota$. At function types we recursively proceed under a semantic λ . The interpretation of contexts is pointwise.

The interpretations of terms and substitutions are coherence properties, which relate the term algebra construction to term evaluation in the $-^A$ model. For terms, if we pick $\nu \equiv \text{id}$, we get $A^T t = t^A \mathbf{TmAlg}_\Omega$. The left side embeds t in the term model via $-^T$, while the right hand side evaluates t in the term model.

A way to view the term algebra construction, is that we are working in a *slice model* over the fixed Ω , and every $\nu : \mathbf{Sub} \Omega \Gamma$ can be viewed as an internal Γ -algebra in this model. The term algebra construction demonstrates that every such internal algebra yields an external element of Γ^A .

2.3.1 Recursor Construction

We show that \mathbf{TmAlg}_Ω supports a recursion principle, i.e. it is weakly initial.

Definition 10 (Recursor construction). We assume $(X, \omega) : \mathbf{Alg} \Omega$; recall that $X : \mathbf{Set}$ and $\omega : \Omega^A X$. We define $\mathbf{R} : \mathbf{T} \rightarrow X$ as $\mathbf{R} t \equiv t^A \omega$. We additionally define the following.

$$\begin{aligned}
& -^R : (A : \mathbf{Ty})(t : \mathbf{Tm} \Omega A) \rightarrow A^M \mathbf{R} (A^T t) (t^A \omega) \\
& \iota^R \quad t \equiv (\text{refl} : t^A \omega = t^A \omega) \\
& (\iota \rightarrow A)^R t \equiv \lambda u. A^R (\text{app } t u)
\end{aligned}$$

$$\begin{aligned} -^R : (\Gamma : \mathbf{Con})(\nu : \mathbf{Sub} \, \Omega \, \Gamma) &\rightarrow \Gamma^M \mathbf{R} \, (\Gamma^T \nu) \, (\nu^A \omega) \\ \Gamma^R \nu x &:= A^R (\nu x) \end{aligned}$$

We define the recursor for Ω as

$$\begin{aligned} \mathbf{Rec}_\Omega : (alg : \mathbf{Alg} \, \Omega) &\rightarrow \mathbf{Mor} \, \mathbf{TmAlg}_\Omega \, alg \\ \mathbf{Rec}_\Omega (X, \omega) &:= (\mathbf{R}, \Omega^R \Omega \, \mathbf{id}) \end{aligned}$$

In short, the way we get recursion is by evaluating terms in arbitrary (X, ω) algebras using $-^A$. The $-^R$ operation for types and contexts confirms that \mathbf{R} preserves structure appropriately, so that \mathbf{R} indeed yields algebra morphisms.

We skip interpreting terms and substitutions by $-^R$. It is necessary to do so with more general signatures, but not in the current chapter.

2.3.2 Eliminator Construction

We take the idea of the previous section a bit further. We have seen that recursion for term algebras is given by evaluation in the “standard” model $-^A$. Now, we show that induction for term algebras corresponds to evaluation in the logical predicate model $-^D$.

Definition 11 (Eliminator construction). We assume $(X^D, \omega^D) : \mathbf{DispAlg} \, \mathbf{TmAlg}_\Omega$. Recall that $X^D : \mathbf{T} \rightarrow \mathbf{Set}$ and $\omega^D : \Omega^D X^D (\Omega^T \Omega \, \mathbf{id})$. Like before, we first interpret the underlying set:

$$\begin{aligned} \mathbf{E} : (t : \mathbf{T}) &\rightarrow X^D t \\ \mathbf{E} t &:= t^D \omega^D \end{aligned}$$

However, this definition is not immediately well-typed, since $t^D \omega^D$ has type $X^D (t^A (\Omega^T \Omega \, \mathbf{id}))$, so we have to show that $t^A (\Omega^T \Omega \, \mathbf{id}) = t$. This equation says that nothing happens if we evaluate a term with type ι in the term model. We get it from the $-^T$ interpretation of terms: $t^T \mathbf{id} : t[\mathbf{id}] = t^A (\Omega^T \Omega \, \mathbf{id})$, and we also

know that $t[\text{id}] = t$. We interpret types and contexts as well:

$$\begin{aligned} -^E : (A : \mathbf{Ty})(t : \mathbf{Tm} \, \Omega \, A) &\rightarrow A^S \, \mathbf{E} \, (t^A \, (\Omega^T \, \Omega \, \text{id})) \, (t^D \, \omega^D) \\ \iota^E \quad t : (t^A \, (\Omega^T \, \Omega \, \text{id}))^D \, \omega^D &= t^D \, \omega^D \\ (\iota \rightarrow A)^E t &\equiv \lambda u. A^E \, (\text{app} \, t \, u) \end{aligned}$$

$$\begin{aligned} -^E : (\Gamma : \mathbf{Con})(\nu : \mathbf{Sub} \, \Omega \, \Gamma) &\rightarrow \Gamma^S \, \mathbf{E} \, (\nu^A \, (\Omega^T \, \Omega \, \text{id})) \, (\nu^D \, \omega^D) \\ \Gamma^E \, \nu \, x &\equiv A^E \, (\nu \, x) \end{aligned}$$

In ι^E we use the same equation as in the definition of \mathbf{E} . In $(\iota \rightarrow A)^E$ the definition is well-typed because of the same equation, but instantiated for the abstracted u term this time. All of this amounts to some additional path induction and transport fiddling in the (intensional) Agda formalization. We get induction for Ω as below.

$$\begin{aligned} \text{Ind}_\Omega : (alg : \mathbf{DispAlg} \, \mathbf{TmAlg}_\Omega) &\rightarrow \mathbf{Section} \, \mathbf{TmAlg}_\Omega \, alg \\ \text{Ind}_\Omega \, (X^D, \omega^D) &\equiv (E, \Omega^E \, \Omega \, \text{id}) \end{aligned}$$

2.4 Discussion

2.4.1 Comparison to F-algebras

A well-known alternative way for treating inductive types is to use certain cocontinuous endofunctors as a more semantic notion of signatures.

For example, single-sorted inductive types can be presented as endofunctors which preserve colimits of some ordinal-indexed chains. For instance, if we have an κ -cocontinuous $F : \mathbb{C} \rightarrow \mathbb{C}$, then algebras are given as $(X : |\mathbb{C}|) \times (\mathbb{C}(F \, X, X))$, morphisms as commuting squares, and Adámek's theorem [AK79] establishes the existence of initial algebras.

An advantage of this approach is that we can describe different classes of signatures by choosing different \mathbb{C} categories:

- If \mathbb{C} is **Set**, we get simple inductive types.
- If \mathbb{C} is **Set** ^{I} for some set I , we get indexed inductive types.
- If \mathbb{C} is **Set**/ I , we get inductive-recursive types.

Another advantage of F -algebras is that signatures are a fairly semantic notion: they make sense even if we have no syntactic presentation at hand. That said, often we do need syntactic signatures, for use in proof assistants, or just to have a convenient notation for a class of cocontinuous functors.

An elegant way of carving out a large class of such functors is to consider polynomials as signatures. For example, when working in **Set**, a signature is an element of $(S : \mathbf{Set}) \times (P : S \rightarrow \mathbf{Set})$, and (S, P) is interpreted as a functor as $X \mapsto (s : S) \times (P s \rightarrow X)$. The initial algebra is the W-type specified by S shapes and P positions. This yields infinitary inductive types as well.

However, it is not known how to get *inductive-inductive* signatures by picking the right \mathbb{C} category and a functor. In an inductive-inductive signature, there may be multiple sorts, which can be indexed over previously declared sorts. For example, in the signature for categories we have $\mathbf{Obj} : \mathbf{Set}$ and $\mathbf{Mor} : \mathbf{Obj} \rightarrow \mathbf{Obj} \rightarrow \mathbf{Set}$, indexed twice over \mathbf{Obj} . Some extensions are required to the idea of F -algebras:

- For inductive-inductive definitions with two sorts, Forsberg gives a specification with two functors, and a considerably more complex notion of algebras, involving dialgebras [NF13]¹.
- For an arbitrary number of sorts, Altenkirch et al. [ACD⁺18] use a “list” of functors, specified mutually with categories of algebras: each functor has as domain the semantic category of all previous sorts.

The functors-as-signatures approach gets significantly less convenient as we consider more general specifications. The approach of this thesis is to skip the middle ground between syntactic signatures and semantic categories of algebras: we treat syntactic signatures as a key component, and give direct semantic interpretation for them. Although we lose the semantic nature of F -algebras, our approach scales extremely well, all the way up to infinitary quotient-inductive-inductive types in Chapter 5, and to some extent to higher inductive-inductive types as well in Chapter 6.

If we look back at $-^A : \mathbf{Con} \rightarrow \mathbf{Set} \rightarrow \mathbf{Set}$, we may note that Γ^A yields a functor, in fact the same functor (up to isomorphism) that we would get from

¹However, the dialgebra specification only covers restricted signatures, where $B : A \rightarrow \mathbf{Set}$ constructor types may refer to $A : \mathbf{Set}$ constructors, but no other dependency is allowed. There is a more general and yet more complicated notion of signature in [NF13], which is not anymore represented with functors.

an F -algebra presentation. However, this is a coincidence in the single-sorted case. With the F -algebra presentation, we can view $(X : |\mathbb{C}|) \times (\mathbb{C}(F X, X))$ as specifying the category of algebras as the total category of a displayed category (by viewing the Σ -type here as taking total categories; a Σ in **Cat**). In our approach, we aim to get the displayed categories directly, without talking about functors.

2.4.2 Generic Programming

Let's consider now our signatures and term algebras in the context of generic programming. This is largely future work, and we don't elaborate it much. But we can draw some preliminary conclusions and make some comparisons.

If a language can formalize inductive signatures and their semantics, that can be viewed as an implementation of generic programming over the described types. Compared to a purely mathematical motivation for this formalization, the requirements for practical generic programming are a bit more stringent.

- *Encoding overhead*: there should be an acceptable overhead in program size and performance when using generic representations. Size blowup can be an issue when writing proofs as well, when types and expressions become too large to mentally parse.
- *Strictness properties*: generic representations should compute as much as possible, ideally in exactly the same way as their non-generic counterparts.

Fixpoints of functors

There is a sizable literature of using fixpoints of functors in generic programming, mainly in Haskell [Swi08, BH11, YHLJ09] and Agda [LM11, AAC⁺20]. We give a minimal example below for an Agda-like implementation.

We have an inductive syntax for some strictly positive functors, covering es-

sentially the same signatures as `Con`.

```

Sig    : Set
Id      : Sig
KT      : Sig
-⊗-    : Sig → Sig → Sig
-⊕-    : Sig → Sig → Sig

```

`Id` codes the identity functor, and `KT` codes the functor which is constantly \top . `-⊗-` and `-⊕-` are pointwise products and coproducts respectively. So we have the evident interpretation functions:

```

[[ - ]] : Sig → Set → Set
map : (F : Sig) → (X0 → X1) → [[F]] X0 → [[F]] X1

```

In Haskell and Agda, the easiest way to get initial algebras is to directly define the inductive fixpoint for each assumed $F : \text{Sig}$:

```

FixF : Set
conF : [[F]] FixF → FixF

```

In Haskell, this definition is valid for arbitrary *semantic* F functor, because there is no termination checking or positivity checking. In Agda, the above definition is valid because the positivity checker is willing to look inside the definition of `[[-]]`. However, Coq and Lean reject this definition. Next we establish weak initiality, by defining the recursor:

```

rec : ([[F]] X → X) → FixF → X
rec f (con x) ≡ f (map (rec f) x)

```

This is again fine in Haskell, but it unfortunately does not pass Agda's termination checker. The most conservative solution is to inline the recursive call into `map`, so that the definition becomes transparent to termination checking.

```

rec : ([[F]] X → X) → FixF → X
rec f (con x) ≡ f (maprec f x)

```

$$\begin{aligned}
\text{maprec} &: \{G\} \rightarrow (\llbracket F \rrbracket X \rightarrow X) \rightarrow \llbracket G \rrbracket \text{Fix}_F \rightarrow \llbracket G \rrbracket X \\
\text{maprec } \{\text{Id}\} \quad f x &:: \equiv \text{rec } f x \\
\text{maprec } \{\text{K}\top\} \quad f x &:: \equiv x \\
\text{maprec } \{G \otimes H\} f (x, y) &:: \equiv (\text{maprec } f x, \text{maprec } f y) \\
\text{maprec } \{G \oplus H\} f (\text{inj}_1 x) &:: \equiv \text{maprec } f x \\
\text{maprec } \{G \oplus H\} f (\text{inj}_2 x) &:: \equiv \text{maprec } f x
\end{aligned}$$

Alternatively, we might use *sized types* [AVW17], as in [AAC⁺20]. The drawback is dependence on an additional language feature which is only supported in Agda, and which has had several soundness issues since its introduction [?].

There is yet another possible approach: defining initial algebras as sequential colimits, using Adámek’s theorem. This approach was taken by Ahrens, Matthes and Mörtberg in [AMM19]. However, the encoding overhead is excessive, and it is practically unusable for generic programming. Another drawback is that defining colimits requires quotient types, which are often not available natively.

W-types

Given a polynomial $(S, P) : (S : \text{Set}) \times (S \rightarrow \text{Set})$, the corresponding W-type is inductively specified as below:

$$\begin{aligned}
W_{S,P} &: \text{Set} \\
\text{sup} &: (s : S) \rightarrow (P s \rightarrow W_{S,P}) \rightarrow W_{S,P}
\end{aligned}$$

If we assume \top , \perp , **Bool**, Π , Σ , W-types, universes and an intensional identity type, a large class of inductive types can be derived, including infinitary and finitary indexed inductive families; this was shown by Hugunin [Hug21]. The encoding also yields definitional β -rules for recursion and elimination. However, there is also significant encoding overhead here.

- First, there is a translation from more convenient signatures to (S, P) polynomials.
- Then, we take the $W_{S,P}$ type, but we need to additionally restrict it to the *canonical* elements by a predicate, as in $(x : W_{S,P}) \times \text{canonical } x$. This is required because the only way to represent inductive branching is by

functions, but functions sometimes contain too many elements up to definitional equality. For example, $\perp \rightarrow A$ has infinitely many definitionally distinct inhabitants.

There is also a performance overhead imposed by the mandatory higher-order constructors. W-types are a great way to have a small basis in a formal setting, both in intensional and extensional type theories, but they are a bit too heavy for practical purposes.

Term algebras

The main advantage of the term algebras described in this chapter is that they can be defined using plain inductive families, which are available in every major dependently typed language. Even in GHC Haskell, the support for generalized ADTs is sufficient for implementing term algebras. In contrast, as we have seen, the direct inductive definition of fixpoints is only possible in Agda, while the colimit definition requires quotients and is rather heavyweight. Even in Agda, the drawback of the direct fixpoint definition is that many generic operations only pass termination checking after manual inlining. With term algebras, generic operations can be defined by induction on $\text{Tm } \Gamma A$ in an obviously terminating way.

For practical usage it makes sense to slightly modify terms. We switch to a *spine neutral* definition. We mutually inductively define **Spine** and **Tm**:

$$\begin{aligned}
 \text{Spine} &: (\Gamma : \text{Con}) \rightarrow \text{Ty} \rightarrow \text{Ty} \rightarrow \text{Set} \\
 \epsilon &: \{A\} \rightarrow \text{Spine } \Gamma A A \\
 \neg, - &: \{B C\} \rightarrow \text{Tm } \Gamma \iota \rightarrow \text{Spine } \Gamma B C \rightarrow \text{Spine } \Gamma (\iota \rightarrow B) C \\
 \text{Tm} &: \text{Con} \rightarrow \text{Ty} \rightarrow \text{Set} \\
 -\$- &: \{A B\} \rightarrow \text{Var } \Gamma A \rightarrow \text{Spine } \Gamma A B \rightarrow \text{Tm } \Gamma B
 \end{aligned}$$

With this representation, a term is always a variable applied to a list of arguments. This can be useful, because pattern matching implementations in metalanguages (e.g. Agda or Idris) are more likely to know about which constructors are possible in patterns. Using this, we give here an ad-hoc definition of natural numbers in pseudo-Agda.

$$\begin{aligned}
 \text{Nat} &: \text{Set} & \text{zero} &:: \equiv \text{vz } \$ \epsilon \\
 \text{Nat} &:: \equiv \text{Tm } (\bullet \triangleright \iota \rightarrow \iota \triangleright \iota) \iota & \text{suc } n &:: \equiv \text{vs vz } \$ (n, \epsilon)
 \end{aligned}$$

$$\begin{aligned}
\text{NatElim} &: (P : \text{Nat} \rightarrow \text{Set}) \rightarrow P \text{ zero} \rightarrow ((n : \text{Nat}) \rightarrow P n \rightarrow P (\text{suc } n)) \\
&\rightarrow (n : \text{Nat}) \rightarrow P n \\
\text{NatElim } P \text{ pz ps } (\text{vz } \$ \epsilon) &\quad : \equiv \text{pz} \\
\text{NatElim } P \text{ pz ps } (\text{vs vz } \$ (n, \epsilon)) &: \equiv \text{ps } n (\text{NatElim } P \text{ pz ps } n)
\end{aligned}$$

The actual Agda definition can be found in the supplementary formalization, and it is pretty much the same as above. We recover the exact same behavior with respect to pattern matching as with native inductive definitions.

Rec_Ω and Ind_Ω can be adapted to spine neutral terms with minor adjustments. But what about the β -rules which they produce as part of their output, are they definitional (i.e. proven by `refl`)? In this chapter we do not have a rigorous way of reasoning about definitional equalities; in the next chapter we develop such reasoning and show that Rec_Ω enjoys definitional β -rules (with or without the spine neutral definition).

However, Ind_Ω only supports propositional β -rules. The issue is the extra transporting in the definition of $\text{E} : (t : \text{Tm } \Omega \iota) \rightarrow X^D t$: transports don't strictly commute with constructors. This appears to be a point of advantage for the direct fixpoint definition in Agda, as it allows generic elimination with strict computation rules [?].

The term algebra presentation can be easily extended to indexed families. In that case, signatures and terms are still definable with basic inductive families, without requiring quotients or complicated encodings; see Kaposi and von Raumer [KvR20].

A related existing work is *sum-of-products* generics by De Vries and Löh [dVL14]. There, signatures for functors are in a normal form: we cannot freely take products and coproducts, instead a signature looks very much like a `Con` in this chapter (except in an indexed form). The authors observe that several generic programming patterns are easier to express with normalized signatures. However, they still use explicit fixpoints as the way to get initial algebras.

CHAPTER 3

Semantics in Two-Level Type Theory

In this chapter we describe how two-level type theory is used as a metatheoretic setting in the rest of this thesis. First, we provide motivation and overview. Second, we describe models of type theories in general, and models of two-level type theories as extensions. Third, we describe presheaf models of two-level type theories. Finally, we generalize the semantics and the term algebra construction from Chapter 2 in two-level type theory, as a way to illustrate the applications.

3.1 Motivation

We note two shortcomings of the semantics presented in the previous chapter.

First, the semantics that we provided was not as general as it could be. We used the internal **Set** universe to specify algebras, but algebras make sense in many different categories. A crude way to generalize semantics is to simply say that our formalization, which was carried out in the syntax (i.e. initial model) of some intensional type theory, can be interpreted in any model of the type theory. However, this is wasteful: for simple inductive signatures, it is enough to assume a category with finite products as semantic setting. We don't need all the extra baggage that comes with a model of a type theory.

Second, we were not able to reason about definitional equalities, only propositional ones. We have a formalization of signatures and semantics in intensional Agda, where the two notions differ¹, but only propositional equality is subject to internal reasoning. For instance, we would like to show that term algebras support recursion with strict β -rules, and for this we need to reason about strict

¹As opposed to in extensional type theory, where they're the same.

equality.

Notation 3. We use \bullet for the terminal object in a \mathbb{C} category, with $\epsilon : \mathbb{C}(A, \bullet)$ for the unique morphism. For products, we use $- \otimes -$ with $(-, -) : \mathbb{C}(A, B) \rightarrow \mathbb{C}(A, C) \rightarrow \mathbb{C}(A, B \otimes C)$ and \mathfrak{p} and \mathfrak{q} for first and second projections respectively.

Example 6. Assuming \mathbb{C} has finite products, natural number algebras and binary tree algebras are specified as follows.

$$\begin{aligned} \mathbf{Alg}_{\mathbf{NatSig}} &:\equiv (X : |\mathbb{C}|) \times \mathbb{C}(\bullet, X) \times \mathbb{C}(X, X) \\ \mathbf{Alg}_{\mathbf{TreeSig}} &:\equiv (X : |\mathbb{C}|) \times \mathbb{C}(\bullet, X) \times \mathbb{C}(X \otimes X, X) \end{aligned}$$

Here, $\mathbf{Alg}_{\mathbf{NatSig}}$ and $\mathbf{Alg}_{\mathbf{TreeSig}}$ are both sets in some metatheory, and the \times in the definitions refer to the metatheoretic Σ . Algebras can be viewed as diagrams which preserve finite products, and algebra morphisms are natural transformations.

How should we adjust \mathbf{Alg} to compute algebras in \mathbb{C} , and \mathbf{Mor} to compute their morphisms? While it is possible to do this in a direct fashion, working directly with objects and morphisms of \mathbb{C} is rather unwieldy. \mathbb{C} is missing many convenience features of type theories.

- There are no variables or binders. We are forced to work in a point-free style or chase diagrams; both become difficult to write and read after a certain point of complexity.
- There are no functions, universes or inductive types.
- Substitution (with weakening as a special case) has to be handled explicitly and manually. Substitutions are certain morphisms, while “terms” are also morphisms, and we have to use composition to substitute terms. In contrast, if we are working internally in a type theory, terms and substitutions are distinct, and we only have to explicitly deal with terms, and substitutions are automated and implicit.

The above overlaps with motivations for working in *internal languages* [nc21] of structured categories: they aid calculation and compact formalization by hiding bureaucratic structural details.

A finite product category \mathbb{C} does not have much of an internal language, it is too bare-bones. But we can work instead in the internal language of $\hat{\mathbb{C}}$, the

category of presheaves over \mathbb{C} . This allows faithful reasoning about \mathbb{C} , while also including all convenience features of extensional type theory.

Two-level type theories [ACKS19], or 2LTT in short, are type theories such that they have “standard” interpretations in presheaf categories. A 2LTT has an inner layer, where types and terms arise by embedding \mathbb{C} in $\hat{\mathbb{C}}$, and an outer layer, where constructions are inherited from $\hat{\mathbb{C}}$. The exact details of the syntax may vary depending on what structures \mathbb{C} supports, and which type formers we assume in the outer layer. Although it is possible to add assumptions to a 2LTT which preclude standard presheaf semantics [ACKS19, Section 2.4.], we stick to basic 2LTT in this thesis. By using 2LTT, we are able to use a type-theoretic syntax which differs only modestly from the style of definitions that we have seen so far.

From a programming perspective, basic 2LTT provides a convenient syntax for writing metaprograms. This can be viewed as *two-stage compilation*: if we have a 2LTT program with an inner type, we can run it, and it returns another program, which lives purely in the inner theory.

3.2 Models of Type Theories

Before explaining 2LTT-specific features, we review models of type theories in general. Variants of 2LTT will be obtained by adding extra features on the top of more conventional TTs.

It is also worth to take a more general look at models at this point, because the notions presented in this subsection (categories with families, type formers) will be reused several times in this thesis, when specifying theories of signatures.

3.2.1 The Algebraic View

We take an algebraic view [?] of models and syntaxes of type theories throughout this thesis. Models of type theories are algebraic structures: they are categories with certain extra structure. The syntax of a type theory is understood to be its initial model. In initial models, the underlying category is the category of typing contexts and parallel substitutions, while the extra structure corresponds to type and term formers, and equations quotient the syntax by definitional equality.

Type theories can be described with quotient inductive-inductive (QII) signatures, and their initial models are quotient inductive-inductive types (QIITs).

Hence, 2LTT is also a QII theory. We will first talk about QIITs in Chapter 4. Until then, we shall make do with an informal understanding of categorical semantics for type theories, without using anything in particular from the metatheory of QIITs. There is some circularity here, that we talk about QIITs in this thesis, but we employ QIITs when talking about them. However, this is only an annoyance in exposition and not a fundamental issue: Chapter ?? describes how to eliminate circularity by a form of bootstrapping.

The algebraic view lets us dispense with all kinds of “raw” syntactic objects. We only ever talk about well-typed and well-formed objects, moreover, every construction must respect definitional equalities. For terms in the algebraic syntax, definitional equality coincides with metatheoretic equality. This mirrors equality of morphisms in 1-category theory, where we usually reuse metatheoretic equality in the same way.

In the following we specify notions of models for type theories. We split this in two parts: categories with families and type formers.

3.2.2 Categories With Families

Definition 12. A **category with family** (cwf) [Dyb95] is a way to specify the basic structural rules for contexts, substitutions, types and terms. It yields a dependently typed explicit substitution calculus [?]. A cwf consists of the following.

- A category with a terminal object. We denote the set of objects as $\mathbf{Con} : \mathbf{Set}$ and use capital Greek letters starting from Γ to refer to objects. The set of morphisms is $\mathbf{Sub} : \mathbf{Con} \rightarrow \mathbf{Con} \rightarrow \mathbf{Set}$, and we use σ, δ and so on to refer to morphisms. We write id for the identity morphism and $- \circ -$ for composition. The terminal object is \bullet with unique morphism $\epsilon : \mathbf{Sub} \Gamma \bullet$. In initial models (that is, syntaxes) of type theories, objects correspond to typing contexts, morphisms to parallel substitutions and the terminal object to the empty context; this informs the naming scheme.
- A *family structure*, containing $\mathbf{Tm} : (\Gamma : \mathbf{Con}) \rightarrow \mathbf{Tm} \Gamma \rightarrow \mathbf{Set}$. We use A, B, C to refer to types and t, u, v to refer to terms. \mathbf{Tm} is a presheaf over the category of contexts and \mathbf{Tm} is a displayed presheaf

over \mathbf{Ty} . This means that types and terms can be substituted:

$$\begin{aligned} -[-] : \mathbf{Ty} \Delta \rightarrow \mathbf{Sub} \Gamma \Delta \rightarrow \mathbf{Ty} \Gamma \\ -[-] : \mathbf{Tm} \Delta A \rightarrow (\sigma : \mathbf{Sub} \Gamma \Delta) \rightarrow \mathbf{Tm} \Gamma (A[\sigma]) \end{aligned}$$

Substitution is functorial: we have $A[\mathbf{id}] = A$ and $A[\sigma \circ \delta] = A[\sigma][\delta]$, and likewise for terms.

A family structure is additionally equipped with *context comprehension* which consists of a context extension operation $- \triangleright - : (\Gamma : \mathbf{Con}) \rightarrow \mathbf{Ty} \Gamma \rightarrow \mathbf{Con}$ together with an isomorphism $\mathbf{Sub} \Gamma (\Delta \triangleright A) \simeq ((\sigma : \mathbf{Sub} \Gamma \Delta) \times \mathbf{Tm} \Gamma (A[\sigma]))$ which is natural in Γ .

The following notions are derivable from the comprehension structure:

- By going right-to-left along the isomorphism, we recover *substitution extension* $-, - : (\sigma : \mathbf{Sub} \Gamma \Delta) \rightarrow \mathbf{Tm} \Gamma (A[\sigma]) \rightarrow \mathbf{Sub} \Gamma (\Delta \triangleright A)$. This means that starting from ϵ or the identity substitution \mathbf{id} , we can iterate $-, -$ to build substitutions as lists of terms.
- By going left-to-right, and starting from $\mathbf{id} : \mathbf{Sub} (\Gamma \triangleright A) (\Gamma \triangleright A)$, we recover the *weakening substitution* $\mathbf{p} : \mathbf{Sub} (\Gamma \triangleright A) \Gamma$ and the *zero variable* $\mathbf{q} : \mathbf{Tm} (\Gamma \triangleright A) (A[\mathbf{p}])$.
- By weakening \mathbf{q} , we recover a notion of variables as De Bruijn indices. In general, the n -th De Bruijn index is defined as $\mathbf{q}[\mathbf{p}^n]$, where \mathbf{p}^n denotes n -fold composition.

Comprehension can be characterized either by taking $-, -, \mathbf{p}$ and \mathbf{q} as primitive, or the natural isomorphism. The two are equivalent, and we may switch between them, depending on which is more convenient.

There are other ways for presenting the basic categorical structure of models, which are nonetheless equivalent to cwfs, including natural models [Awo18] and categories with attributes [Car78]. We use the cwf presentation for its immediately algebraic character and closeness to conventional explicit substitution syntax.

Notation 4. As De Bruijn indices are hard to read, we will mostly use nameful notation for binders. For example, assuming $\mathbf{Nat} : \{\Gamma : \mathbf{Con}\} \rightarrow \mathbf{Ty} \Gamma$ and $\mathbf{Id} : \{\Gamma : \mathbf{Con}\}(A : \mathbf{Ty} \Gamma) \rightarrow \mathbf{Tm} \Gamma A \rightarrow \mathbf{Tm} \Gamma A \rightarrow \mathbf{Ty} \Gamma$, we may write $\bullet \triangleright n : \mathbf{Nat} \triangleright$

$p : \text{Id Nat } n \ n$ for a typing context, instead of using numbered variables or cwf combinators as in $\bullet \triangleright \text{Nat} \triangleright \text{Id Nat } q \ q$.

Notation 5. In the following, we will denote family structures by (Ty, Tm) pairs and overload context extension $- \triangleright -$ for different families.

Definition 13. The following derivable operations are commonly used.

- *Single substitution* can be derived from parallel substitution as follows. Assume $t : \text{Tm } (\Gamma \triangleright A) \ B$, and $u : \text{Tm } \Gamma \ A$. t is a term which may depend on the last variable in the context, which has A type. We can substitute that variable with the u term as $t[\text{id}, u] : \text{Tm } \Gamma \ (A[\text{id}, u])$. Note that term substitution causes the type to be substituted as well. $(\text{id}, u) : \text{Sub } \Gamma \ (\Gamma \triangleright A)$ is well-typed because $u : \text{Tm } \Gamma \ A$ hence also $u : \text{Tm } \Gamma \ (A[\text{id}])$.
- We can *lift substitutions* over binders as follows. Assuming $\sigma : \text{Sub } \Gamma \ \Delta$ and $A : \text{Ty } \Delta$, we construct a lifting of σ which maps an additional A -variable to itself: $(\sigma \circ \mathbf{p}, \mathbf{q}) : \text{Sub } (\Gamma \triangleright A[\sigma]) \ (\Delta \triangleright A)$. Let us see why this is well-typed. We have $\mathbf{p} : \text{Sub } (\Gamma \triangleright A[\sigma]) \ \Gamma$ and $\sigma : \text{Sub } \Gamma \ \Delta$, so $\sigma \circ \mathbf{p} : \text{Sub } (\Gamma \triangleright A[\sigma]) \ \Delta$. Also, $\mathbf{q} : \text{Tm } (\Gamma \triangleright A[\sigma]) \ (A[\sigma][\mathbf{p}])$, hence $\mathbf{q} : \text{Tm } (\Gamma \triangleright A[\sigma]) \ (A[\sigma \circ \mathbf{p}])$, thus $(\sigma \circ \mathbf{p}, \mathbf{q})$ typechecks.

Notation 6. As a nameful notation for substitutions, we may write $t[x \mapsto u]$, for a single substitution, or $t[x \mapsto u_1, y \mapsto u_2]$ and so on.

In nameful notation we leave all weakening implicit, including substitution lifting. Formally, if we have $t : \text{Tm } \Gamma \ A$, we can only mention t in Γ . If we need to mention it in $\Gamma \triangleright B$, we need to use $t[\mathbf{p}]$ instead. In the nameful notation, $t : \text{Tm } (\Gamma \triangleright x : B) \ A$ may be used.²

3.2.3 Type formers

A family structure in a cwf may be closed under certain type formers, such as functions, Σ -types, universes or inductive types. We give some examples here for their specification. First, we look at common negative type formers, which can be specified using isomorphisms. Then, we consider positive type formers, and finally universes.

²Moreover, when working in the internal syntax of a theory, we just write Agda-like type-theoretic notation, without noting contexts and substitutions in any way.

Negative types

Definition 14. A $(\mathsf{T}_y, \mathsf{T}_m)$ family supports **Π -types** if it supports the following.

$$\begin{aligned}
\Pi & : (A : \mathsf{T}_y \Gamma) \rightarrow \mathsf{T}_y (\Gamma \triangleright A) \rightarrow \mathsf{T}_y \Gamma \\
\Pi[] & : (\Pi A B)[\sigma] = \Pi (A[\sigma]) (B[\sigma \circ \mathbf{p}, \mathbf{q}]) \\
\mathsf{app} & : \mathsf{T}_m \Gamma (\Pi A B) \rightarrow \mathsf{T}_m (\Gamma \triangleright A) B \\
\mathsf{lam} & : \mathsf{T}_m (\Gamma \triangleright A) B \rightarrow \mathsf{T}_m \Gamma (\Pi A B) \\
\Pi\beta & : \mathsf{app} (\mathsf{lam} t) = t \\
\Pi\eta & : \mathsf{lam} (\mathsf{app} t) = t \\
\mathsf{lam}[] & : (\mathsf{lam} t)[\sigma] = \mathsf{lam} (t[\sigma \circ \mathbf{p}, \mathbf{q}])
\end{aligned}$$

Here, Π is the type formation rule. $\Pi[]$ is the type substitution rule, expressing that substituting Π proceeds structurally on constituent types. Note $B[\sigma \circ \mathbf{p}, \mathbf{q}]$, where we lift σ over the additional binder.

The rest of the rules specify a natural isomorphism $\mathsf{T}_m \Gamma (\Pi A B) \simeq \mathsf{T}_m (\Gamma \triangleright A) B$. We only need a substitution rule (i.e. a naturality rule) for one direction of the isomorphism, since the naturality of the other map is derivable.

This way of specifying Π -types is very convenient if we have explicit substitutions. The usual “pointful” specification is equivalent to this. For example, we have the following derivation of pointful application:

$$\begin{aligned}
\mathsf{app}' & : \mathsf{T}_m \Gamma (\Pi A B) \rightarrow (u : \mathsf{T}_m \Gamma A) \rightarrow \mathsf{T}_m \Gamma (B[\mathbf{id}, u]) \\
\mathsf{app}' t u & \equiv (\mathsf{app} t)[\mathbf{id}, u]
\end{aligned}$$

Remark on naturality. The above specification for Π can be written more compactly if we assume that everything is natural with respect to substitution.

$$\begin{aligned}
\Pi & : (A : \mathsf{T}_y \Gamma) \rightarrow \mathsf{T}_y (\Gamma \triangleright A) \rightarrow \mathsf{T}_y \Gamma \\
(\mathsf{app}, \mathsf{lam}) & : \mathsf{T}_m \Gamma (\Pi A B) \simeq \mathsf{T}_m (\Gamma \triangleright A) B
\end{aligned}$$

This is a reasonable assumption; in the rest of the thesis we only ever define structures on cwfs which are natural in this way.

Notation 7. From now on, when specifying type formers in family structures, we assume that everything is natural, and thus omit substitution equations.

There are ways to make this idea more precise, and take it a step further by working in languages where only natural constructions are possible. The term

higher-order abstract syntax (HOAS) is sometimes used for this style. It lets us also omit contexts, so we would only need to write

$$\begin{aligned} \Pi & : (A : \mathbf{Ty}) \rightarrow (\mathbf{Tm} A \rightarrow \mathbf{Ty}) \rightarrow \mathbf{Ty} \\ (\mathbf{app}, \mathbf{lam}) & : \mathbf{Tm} (\Pi A B) \simeq ((a : \mathbf{Tm} A) \rightarrow \mathbf{Tm} (B a)) \end{aligned}$$

Recently several promising works emerged in this area [?]. Although this technology is likely to be the preferred future direction in the metatheory of type theories, this thesis does not make use of it. The field is rather fresh, with several different approaches and limited amount of pedagogical exposition, and the new techniques would also raise the level of abstraction in this thesis, all contributing to making it less accessible. It's also not obvious how exactly HOAS-style could be employed to aid formalization here, and it would require significant additional research. Often, a setup with multiple modalities (“multimodal” [GKNB20]) is required [BKS21], because we work with presheaves over different cwfs. It seems that a synthetic notion of dependent modes would be also required to formalize constructions in this thesis, since we often work with displayed presheaves over displayed cwfs. This is however not yet developed in literature.

Definition 15. A family structure supports **constant families** if we have the following.

$$\begin{aligned} \mathbf{K} & : \mathbf{Con} \rightarrow \{\Gamma : \mathbf{Con}\} \rightarrow \mathbf{Ty} \Gamma \\ (\mathbf{app}_\mathbf{K}, \mathbf{lam}_\mathbf{K}) & : \mathbf{Tm} \Gamma (\mathbf{K} \Delta) \simeq \mathbf{Sub} \Gamma \Delta \end{aligned}$$

Constant families express that every context can be viewed as a non-dependent type in any context. Having constant families is equivalent to the *democracy* property for a cwf [CD14, NF13]. Constant families are convenient when building models, because they let us model non-dependent types as semantic contexts, which are often simpler structures than semantic types. From a programming perspective, constant families specify closed record types, where $\mathbf{K} \Delta$ has Δ -many fields.

If we have equalities of sets for the specification, i.e. $\mathbf{Tm} \Gamma (\mathbf{K} \Delta) = \mathbf{Sub} \Gamma \Delta$, we have **strict constant families**.

Definition 16. A family structure supports **Σ -types** if we have

$$\begin{aligned} \Sigma & : (A : \mathbf{Ty} \Gamma) \rightarrow \mathbf{Ty} (\Gamma \triangleright A) \rightarrow \mathbf{Ty} \Gamma \\ (\mathbf{proj}, (-, -)) & : \mathbf{Tm} \Gamma (\Sigma A B) \simeq ((t : \mathbf{Tm} \Gamma A) \times \mathbf{Tm} \Gamma (B[\mathbf{id}, t])) \end{aligned}$$

We use the shorter specification above, where everything is assumed to be natural. We may write proj_1 and proj_2 for composing the metatheoretic first and second projections with proj .

Definition 17. A family structure supports the **unit type** if we have $\top : \text{Ty } \Gamma$ such that $\text{Tm } \Gamma \top \simeq \top$, where the \top on the right is the metatheoretic unit type, and we overload \top for the internal unit type. From this, we get the internal $\text{tt} : \text{Tm } \Gamma \top$, which is definitionally unique.

Definition 18. A family structure supports **extensional identity** types if there is $\text{ld} : \text{Tm } \Gamma A \rightarrow \text{Tm } \Gamma A \rightarrow \text{Ty } \Gamma$ such that $(\text{reflect}, \text{refl}) : \text{Tm } \Gamma (\text{ld } t u) \simeq (t = u)$.

It is also possible to give a positive definition for identity types, in which case we get intensional identity. Extensional identity corresponds to a categorical equalizer of terms (a limit), while the Martin-Löf-style intensional identity is characterized as the initial reflexive relation on a type (a colimit).

This choice between negative and positive specification generally exists for type formers with a single term construction rule. For example, Σ can be defined as a positive type, with an elimination rule that behaves like pattern matching. Positive Σ is equivalent to negative Σ , although it only supports propositional η -rules. In contrast, positive identity is usually *not* equivalent to negative identity.

$\text{refl} : t = u \rightarrow \text{Tm } \Gamma (\text{ld } t u)$ expresses reflexivity of identity: definitionally equal terms are provably equal. reflect , which goes the other way around, is called *equality reflection* [?]: provably equal terms are identified in the metatheory.

Uniqueness of identity proofs (UIP) is often ascribed to the extensional identity type [?]. UIP means that $\text{Tm } \Gamma (\text{ld } t u)$ has at most a single inhabitant up to ld . However, UIP is not something which is inherent in the negative specification, instead it's inherited from the metatheory. If Tm forms a homotopy set in the metatheory, then internal equality proofs inherit uniqueness through the defining isomorphism.

Positive types

We do not dwell much on positive types here, as elsewhere in this thesis we talk a lot about specifying such types anyway. We provide here some background and a small example.

The motivation is to specify initial internal algebras in a cwf. However, specifying the uniqueness of recursors using definitional equality is problematic, if we

are to have decidable and efficient conversion checking for a type theory. Consider the specification of **Bool** together with its recursor.

$$\begin{aligned}
\mathbf{Bool} & : \mathbf{Ty} \Gamma \\
\mathbf{true} & : \mathbf{Tm} \Gamma \mathbf{Bool} \\
\mathbf{false} & : \mathbf{Tm} \Gamma \mathbf{Bool} \\
\mathbf{BoolRec} & : (B : \mathbf{Ty} \Gamma) \rightarrow \mathbf{Tm} \Gamma B \rightarrow \mathbf{Tm} \Gamma B \rightarrow \mathbf{Tm} \Gamma \mathbf{Bool} \rightarrow \mathbf{Tm} \Gamma B \\
\mathbf{true}\beta & : \mathbf{BoolRec} B t f \mathbf{true} = t \\
\mathbf{false}\beta & : \mathbf{BoolRec} B t f \mathbf{false} = f
\end{aligned}$$

BoolRec together with the β -rules specifies an internal **Bool**-algebra morphism. A possible way to specify definitional uniqueness is as follows. Assuming $B : \mathbf{Ty} \Gamma$, $t : \mathbf{Tm} \Gamma B$, $f : \mathbf{Tm} \Gamma B$ and $m : \mathbf{Tm} (\Gamma \triangleright b : \mathbf{Bool}) B$, such that $m[b \mapsto \mathbf{true}] = t$ and $m[b \mapsto \mathbf{false}] = f$, it follows that $\mathbf{BoolRec} B t f b : \mathbf{Tm} (\Gamma \triangleright b : \mathbf{Bool}) B$ is equal to m .

Unfortunately, deciding conversion with this rule entails deciding pointwise equality of arbitrary **Bool** functions, which can be done in exponential time in the number of **Bool** arguments. More generally, Scherer presented a decision algorithm for conversion checking with strong finite sums and products in simple type theory [Sch17], which also takes exponential time. If we move to natural numbers with definitionally unique recursion, conversion checking becomes undecidable.

One solution is to have propositionally unique recursion instead. However, if such equations are postulated, that would break the canonicity property in intensional type theories, since now we would have equality proofs other than **refl** in the empty context.

The standard solution is to have dependent elimination principles instead: this allows inductive reasoning, canonicity and effectively decidable definitional equality at the same time. For **Bool**, we would have

$$\begin{aligned}
\mathbf{BoolInd} & : (B : \mathbf{Ty} (\Gamma \triangleright b : \mathbf{Bool})) \rightarrow \mathbf{Tm} \Gamma (B[b \mapsto \mathbf{true}]) \\
& \rightarrow \mathbf{Tm} \Gamma (B[b \mapsto \mathbf{false}]) \rightarrow (t : \mathbf{Tm} \Gamma \mathbf{Bool}) \rightarrow \mathbf{Tm} \Gamma (B[b \mapsto t])
\end{aligned}$$

together with $\mathbf{BoolInd} B t f \mathbf{true} = t$ and $\mathbf{BoolInd} B t f \mathbf{false} = f$.

Of course, if we assume extensional identity types, we have undecidable conversion anyway, and definitionally unique recursion is equivalent to induction. But decidable conversion is an essential part of type theory, perhaps its main

selling point as a foundation for mechanized mathematics, which makes it possible to relegate a deluge of boilerplate to computers. Hence, decidable conversion should be kept in mind.

Universes

Universes are types which classify types. There are several different flavors of universes.

Definition 19. A **Tarski-style** universe consists of the following data:

$$U : \text{Ty } \Gamma \quad \text{El} : \text{Tm } \Gamma \, U \rightarrow \text{Ty } \Gamma$$

This is a weak classifier, since not all $\text{Ty } \Gamma$ are necessarily represented as terms of the universe. Rather, this kind of universe can be viewed as an internal subfamily of (Ty, Tm) . Like families, Tarski universes can be closed under type formers. For instance, if U has **Nat**, we have the following:

$$\text{Nat} : \text{Tm } \Gamma \, U \quad \text{zero} : \text{Tm } \Gamma \, (\text{El Nat}) \quad \text{suc} : \text{Tm } \Gamma \, (\text{El Nat}) \rightarrow \text{Tm } \Gamma \, (\text{El Nat})$$

$$\begin{aligned} \text{NatElim} : & (P : \text{Ty } (\Gamma \triangleright n : \text{El Nat})) \\ & \rightarrow \text{Tm } \Gamma \, (P[n \mapsto \text{zero}]) \\ & \rightarrow \text{Tm } (\Gamma \triangleright n : \text{El Nat} \triangleright np : P[n \mapsto n]) \, (P[n \mapsto \text{suc } n]) \\ & \rightarrow (n : \text{Tm } \Gamma \, (\text{El Nat})) \rightarrow \text{Tm } \Gamma \, (P[n \mapsto n]) \end{aligned}$$

If all type formers in U follow this scheme, U may be called a **weakly Tarski** universe. If we assume that every type former in U is also duplicated in (Ty, Tm) , moreover El preserves all type formers, so that e.g. El Nat is definitionally equal to the natural number type in Ty , then U is **strongly Tarski**.

It is often more convenient to have stronger classifiers as universes, so that *all* types in a given family structure are represented.

Definition 20. Ignoring size issues for now, **Coquand universes** [?] are specified as follows:

$$U : \text{Ty } \Gamma \quad (\text{El}, \text{c}) : \text{Tm } \Gamma \, U \simeq \text{Ty } \Gamma$$

c maps every type in Ty to a code in U . Now we can ignore El when specifying type formers, as c can be always used to get a code in U for a type.

Unfortunately, the exact specification above yields an inconsistent “type-in-type” system, because \mathbf{U} itself has a code in \mathbf{U} . The standard solution is to have multiple family structures $(\mathbf{T}y_i, \mathbf{T}m_i)$, indexed by universe levels, and have $\mathbf{U}_i : \mathbf{T}y_{i+1} \Gamma$ and $\mathbf{T}m_{i+1} \Gamma \mathbf{U}_i \simeq \mathbf{T}y_i \Gamma$. For a general specification of consistent universe hierarchies, see [Kov21]. As mentioned in Section 1.3.2, we omit universe indices in the following, and implicitly assume “just enough” universes for particular purposes.

Definition 21. **Russell universes** are Coquand universes additionally satisfying $\mathbf{T}m \Gamma \mathbf{U} = \mathbf{T}y \Gamma$ as an equality of sets, and also $\mathbf{E}l t = t$. This justifies omitting $\mathbf{E}l$ and \mathbf{c} from informal notation, implicitly casting between $\mathbf{T}m \Gamma \mathbf{U}$ and $\mathbf{T}y \Gamma$.

Russell-style universes are commonly supported in set-theoretic models. They are also often inherited from meta-type-theories which themselves have Russell-universes. Major implementations of type theories (Coq, Lean, Agda, Idris) are all such.

3.3 Two-Level Type Theory

3.3.1 Models

We describe models of 2LTT in the following. This is not the only possible way to present 2LTT; our approach differs from [ACKS19] in some ways. We will summarize the differences at the end of this section.

Definition 22. A model of a **two-level type theory** is a model of type theory such that

- It supports a Tarski-style universe $\mathbf{T}y_0 : \mathbf{T}y \Gamma$ with decoding $\mathbf{T}m_0 : \mathbf{T}m \Gamma \mathbf{T}y_0 \rightarrow \mathbf{T}y \Gamma$.
- $\mathbf{T}y_0$ may be closed under arbitrary type formers, however, it is only possible to eliminate from $\mathbf{T}y_0$ type formers to types in $\mathbf{T}y_0$.

Types in $\mathbf{T}y_0$ are called *inner* types, while other types are *outer*. Alternatively, we may talk about *object-level* and *meta-level* types.

For example, if we have inner functions, we have the following:

$$\begin{aligned} \Pi_0 & : (A : \mathbf{T}m \Gamma \mathbf{T}y_0) \rightarrow \mathbf{T}m (\Gamma \triangleright \mathbf{T}m_0 A) \rightarrow \mathbf{T}m \Gamma \mathbf{T}y_0 \\ (\mathbf{app}_0, \mathbf{lam}_0) & : \mathbf{T}m \Gamma (\mathbf{T}m_0 (\Pi_0 A B)) \simeq \mathbf{T}m (\Gamma \triangleright \mathbf{T}m_0 A) (\mathbf{T}m_0 B) \end{aligned}$$

If we have inner Booleans, we have the following (with β -rules omitted):

$$\begin{aligned}
\text{Bool}_0 & : \mathsf{Tm} \Gamma \mathsf{T}_{\mathbf{y}_0} \\
\text{true}_0 & : \mathsf{Tm} \Gamma (\mathsf{Tm}_0 \text{Bool}_0) \\
\text{false}_0 & : \mathsf{Tm} \Gamma (\mathsf{Tm}_0 \text{Bool}_0) \\
\text{BoolInd}_0 & : (B : \mathsf{Tm} (\Gamma \triangleright b : \mathsf{Tm}_0 \text{Bool}_0) \mathsf{T}_{\mathbf{y}_0}) \\
& \rightarrow \mathsf{Tm} \Gamma (\mathsf{Tm}_0 (B[b \mapsto \text{true}_0])) \\
& \rightarrow \mathsf{Tm} \Gamma (\mathsf{Tm}_0 (B[b \mapsto \text{false}_0])) \\
& \rightarrow (t : \mathsf{Tm} \Gamma (\mathsf{Tm}_0 \text{Bool}_0)) \rightarrow \mathsf{Tm} \Gamma (\mathsf{Tm}_0 (B[b \mapsto t]))
\end{aligned}$$

Intuitively, we can view outer types and terms as metatheoretical, while $\mathsf{T}_{\mathbf{y}_0}$ represents the set of types in the object theory, and Tm_0 witnesses that any object type can be mapped to a metatheoretical set of object terms. The restriction on elimination is crucial. If we have a Boolean term in the object language, we can use the object-level elimination principle to construct new object terms. But it makes no sense to eliminate into the metatheory. In fact, an object-level Boolean term is not necessarily **true** or **false**, it can also be just a variable or neutral term in some context, or it can be an arbitrary non-canonical value in a given model.

We review some properties of 2LTT. An important point is the action of Tm_0 on type formers. In general, Tm_0 preserves the negative type formers but not others.

For example, we have the isomorphism $\mathsf{Tm}_0 (\Pi_0 A B) \simeq \Pi_1 (\mathsf{Tm}_0 A) (\mathsf{Tm}_0 B)$, where Π_1 denotes outer functions. We move left-to-right by mapping t to $\text{lam}_1 (\text{app}_1 t)$, and the other way by mapping t to $\text{lam}_0 (\text{app}_0 t)$. The preservation of Σ , \top , K and extensional identity is analogous.

In contrast, we can map from outer positive types to inner ones, but not the other way around. From $b : \mathsf{Tm} \Gamma \text{Bool}_1$, we can use the outer Bool_1 recursor to return in $\mathsf{Tm}_0 \text{Bool}_0$. In the other direction, we only have constant functions since the Bool_0 recursor only targets types in $\mathsf{T}_{\mathbf{y}_0}$.

It may be the case that there are universes in the inner layer. For example, disregarding size issues (or just accepting an inconsistent inner theory), there may be an U_0 in $\mathsf{T}_{\mathbf{y}_0}$ such that we have $\mathsf{Tm} \Gamma (\mathsf{Tm}_0 \mathsf{U}_0) = \mathsf{Tm} \Gamma \mathsf{T}_{\mathbf{y}_0}$. This amounts to having a Russell-style inner universe with type-in-type. Assume that we have U_1 as well, as a meta-level Russell universe. Then we can map from $\mathsf{Tm}_0 \mathsf{U}_0$ to U_1 , by taking A to $\mathsf{Tm}_0 A$, but we cannot map in the other direction.

3.3.2 Internal Syntax and Notation

In the rest of this thesis we will mostly work internally to a 2LTT, i.e. we use 2LTT as metatheory. We adapt the metatheoretical notations used up until now. We list used features and conventions below.

- We keep previous notation for type formers. For instance, Π -types are written as $(x : A) \rightarrow B$ or as $A \rightarrow B$.
- We assume a Coquand-style universe in the outer layer, named **Set**. As before, we leave the sizing levels implicit; if we were fully precise, we would write Set_i for a hierarchy of outer universes. Despite having a Coquand universe, we shall omit encoding and decoding in the internal syntax, and instead work in Russell-style. In practical implementations, elaborating Russell-style notation to Coquand-style is straightforward to do.
- If the same type formers are supported both in the inner and outer layers, we may distinguish them by $_0$ and $_1$ subscripts, e.g. by having Bool_0 and Bool_1 . We omit some inferable subscripts, e.g. for Π and Σ -types. In these cases, we usually know from the type parameters which type former is meant. For example, $\text{Tm}_0 \text{Bool}_0 \rightarrow \text{Bool}_1$ can only refer to outer functions.
- We have the convention that $- = -$ refers to the inner equality type, while $- \equiv -$ refers to the outer equality type. In the semantics of 2LTT that we use, outer equality of inner values is interpreted as definitional equality of inner terms, hence the naming scheme.
- By having **Set**, we are able to have $\text{Ty}_0 : \text{Set}$ and $\text{Tm}_0 : \text{Ty}_0 \rightarrow \text{Set}$. So we don't have to deal with proper meta-level types, and have a more uniform notation. Notation and specification for inner type formers changes accordingly. For example, for inner Π -types we may write $(x : A) \rightarrow B$ if $A : \text{Ty}_0$ and B depends on $x : \text{Tm}_0 A$. This also enables a higher-order specification: if $B : \text{Tm}_0 A \rightarrow \text{Ty}_0$, then $(x : A) \rightarrow B x : \text{Ty}_0$, and the specifying isomorphism for Π can be written as $\text{Tm}_0((x : A) \rightarrow B x) \simeq ((x : \text{Tm}_0 A) \rightarrow \text{Tm}_0(B x))$.

Notation 8. An explicit notation for inner function abstraction would look like $\text{lam}_0 t$ for $t : (x : \text{Tm}_0 A) \rightarrow \text{Tm}_0(B x)$. This results in “double” abstraction, e.g. in $\text{lam}_0(\lambda x. \text{suc}_0(\text{suc}_0 x)) : \text{Tm}_0(\text{Nat}_0 \rightarrow \text{Nat}_0)$. Instead of

this, we write $\lambda_0 x. t$ as a notation, thus we write $\lambda_0 x. \text{suc}_0 (\text{suc}_0 x)$ for the above example. We may also group multiple λ_0 binders together the same way as with λ .

- We may omit inferable Tm_0 applications. For instance, $\text{Bool}_1 \rightarrow \text{Bool}_0$ can be “elaborated” to $\text{Bool}_1 \rightarrow \text{Tm}_0 \text{Bool}_0$ without ambiguity, since the function codomain must be on the same level as the domain, and the only thing we can do to make sense of this is to lift the codomain by Tm_0 . Sometimes there is some ambiguity: $(\text{Bool}_0 \rightarrow \text{Bool}_0) \rightarrow \text{Bool}_1$ can be elaborated both to $\text{Tm}_0 (\text{Bool}_0 \rightarrow \text{Bool}_0) \rightarrow \text{Bool}_1$ and to $(\text{Tm}_0 \text{Bool}_0 \rightarrow \text{Tm}_0 \text{Bool}_0) \rightarrow \text{Bool}_1$. However, in this case the two output types are definitionally isomorphic, because of the Π -preservation by Tm_0 . Hence, the elaboration choice does not make much difference, so we may still omit Tm_0 -s in situations like this.

Example 7. Working in the internal syntax of 2LTT, the specification of Bool_0 looks like the following (omitting β again):

$$\begin{aligned} \text{Bool}_0 & : \text{Ty}_0 \\ \text{true}_0 & : \text{Bool}_0 \\ \text{false}_0 & : \text{Bool}_0 \\ \text{BoolInd}_0 & : (B : \text{Bool}_0 \rightarrow \text{Ty}_0) \rightarrow B \text{true}_0 \rightarrow B \text{false}_0 \rightarrow (t : \text{Bool}_0) \rightarrow B t \end{aligned}$$

If we elaborate the type of BoolInd_0 , we get the following:

$$\begin{aligned} \text{BoolInd}_0 & : (B : \text{Tm}_0 \text{Bool}_0 \rightarrow \text{Ty}_0) \rightarrow \text{Tm}_0 (B \text{true}_0) \rightarrow \text{Tm}_0 (B \text{false}_0) \\ & \rightarrow (t : \text{Tm}_0 \text{Bool}_0) \rightarrow \text{Tm}_0 (B t) \end{aligned}$$

Here, the type is forced to live in the outer level because of the dependency on Ty_0 . Since Ty_0 is an outer type, $\text{Bool}_0 \rightarrow \text{Ty}_0$ must be lifted, which in turn requires all other types to be lifted as well.

3.3.3 Alternative Presentation for 2LTT

We digress a bit on a different way to present 2LTT. In the primary 2LTT reference [ACKS19], inner and outer layers are specified as follows. We have two different *family structures* on the base cw_f , $(\text{Ty}_0, \text{Tm}_0)$ and $(\text{Ty}_1, \text{Tm}_1)$, and a morphism between them. A family morphism is natural transformation mapping

types to types and terms to terms, which is an isomorphism on terms. We might name the component maps as follows:

$$\begin{aligned}\Uparrow &: \mathsf{Ty}_0 \Gamma \rightarrow \mathsf{Ty}_1 \Gamma \\ \uparrow &: \mathsf{Tm}_0 \Gamma A \rightarrow \mathsf{Tm}_1 \Gamma (\Uparrow A) \\ \downarrow &: \mathsf{Tm}_1 \Gamma (\Uparrow A) \rightarrow \mathsf{Tm}_0 \Gamma A\end{aligned}$$

An advantage of this presentation is that we may close $(\mathsf{Ty}_0, \mathsf{Tm}_0)$ under type formers without any encoding overhead, for example by having $\mathsf{Bool}_0 : \mathsf{Ty}_0 \Gamma$, $\mathsf{true}_0 : \mathsf{Tm}_0 \Gamma \mathsf{Bool}_0$, etc., without the Tarski-style decoding. On the other hand, we don't automatically get an outer universe of inner types. We can recover that in two ways:

- We can assume an inner universe $\mathsf{U}_0 : \mathsf{Ty}_0 \Gamma$, which can be lifted to the outer theory as $\Uparrow \mathsf{U}_0$. However, we may not want to make this assumption, in order to keep the inner theory as simple as possible.
- We can assume an outer universe which classifies elements of $\mathsf{Ty}_0 \Gamma$. This amounts to reproducing the Ty_0 *type* from our 2LTT presentation, as an additional assumption. But in this case, we might as well skip the two family structures and the \Uparrow morphism.

In this thesis we make ubiquitous use of the outer universe of inner types, so we choose that to be the primitive notion, instead of having two family structures.

Do we lose anything by this? For the purposes of this thesis, not really. However, if we want to implement 2LTT as a system for two-stage compilation, the \Uparrow syntax appears to be closer to existing systems. Staging is about computing all outer redexes but no inner ones, thereby outputting syntax which is purely in the inner theory. This can be implemented as a stage-aware variant of normalization-by-evaluation [?]. We can give an intuitive staging interpretation for the operators in the \Uparrow syntax:

- $\Uparrow A$ is the type of A -expressions. This corresponds to *a code* in MetaOcaml [Kis14] and $\mathsf{TExp} a$ in typed Template Haskell.
- \uparrow is *quoting*, which creates an expression from any inner term. This is $\langle - \rangle$ in MetaOCaml and $[[-]]$ in typed Template Haskell.

- \downarrow is *splicing*, which inserts the result of a meta-level computation into an object-level expression. This is $\sim(-)$ in MetaOCaml and $\$(-)$ in typed Template Haskell.

For example, in the \uparrow syntax, we might write a polymorphic identity function which acts on inner types in two different ways:

$$\begin{aligned} \text{id} &: (A : \mathbf{U}_0) \rightarrow A \rightarrow A & \text{id}' &: (A : \uparrow \mathbf{U}_0) \rightarrow \uparrow(\downarrow A) \rightarrow \uparrow(\downarrow A) \\ \text{id} &\equiv \lambda_0 A x. x & \text{id}' &\equiv \lambda_1 A x. x \end{aligned}$$

The first one lives in the inner family structure. The second one is the same thing, but lifted to the outer theory. The choice between the two allows us to control staging-time evaluation. If we write $\text{id } \mathbf{Bool}_0 \text{ true}_0$, that's an inner expression which goes into the staging output as it is. On the other hand, $\downarrow(\text{id}'(\uparrow \mathbf{Bool}_0)(\uparrow \text{true}_0))$ reduces to $\downarrow(\uparrow \text{true}_0)$ which in turn reduces to true_0 . The same choice can be expressed in our syntax as well:

$$\begin{aligned} \text{id} &: \mathbf{Tm}_0((A : \mathbf{U}_0) \rightarrow A \rightarrow A) & \text{id}' &: (A : \mathbf{Tm}_0 \mathbf{U}_0) \rightarrow \mathbf{Tm}_0 A \rightarrow \mathbf{Tm}_0 A \\ \text{id} &\equiv \lambda_0 A x. x & \text{id}' &\equiv \lambda A x. x \end{aligned}$$

It remains to be checked which style is preferable in a staging implementation. In the \uparrow style, the quoting and splicing operations add noise to core syntax, but they are also mostly inferable during elaboration, and they pack stage-changing information into \uparrow and \downarrow , thereby making it feasible to omit stage annotations in other places in the core syntax. In the \mathbf{T}_0 style, we don't have quote/splice, but we have to keep track of stages in all type/term formers. The author of this thesis has made a prototype implementation of staging in the \uparrow style but not in the \mathbf{T}_0 style [?]. It would be interesting to compare the two in future work.

3.4 Standard Semantics of 2LTT

We review the standard semantics of 2LTT which we use in the rest of the thesis. This justifies the metaprogramming view, that 2LTT allows meta-level reasoning about an inner theory.

We present it two steps, by assuming progressively more structure in the inner theory. First, we only assume a category. This already lets us present a presheaf semantics for the outer layer. Then, we assume a cwf as the inner theory, which lets us interpret \mathbf{T}_0 and \mathbf{Tm}_0 and also consider inner type formers.

3.4.1 Presheaf Model of the Outer Layer

In this subsection we present a presheaf model for the outer layer of 2LTT, that is, the base category together with the terminal object, the $(\mathbf{Ty}, \mathbf{Tm})$ family and some type formers. This presheaf semantics is well-known in the literature [?]. We give a specification which follows [?] and [?] most closely.

In the following, we work outside 2LTT (since we are defining a model of 2LTT), in a suitable metatheory; an extensional type theory with enough **Set** universes suffices.

We assume a \mathbb{C} category. We write $i, j, k : |\mathbb{C}|$ for objects and $f, g, h : \mathbb{C}(i, j)$ for morphisms. We use a different notation than for cwfs before, in order to disambiguate components in \mathbb{C} from components in the presheaf model of 2LTT. We use $\hat{\mathbb{C}}$ to refer to the model which is being defined. We use the same component names for $\hat{\mathbb{C}}$ as in Section 3.2.

Model of cwf

Definition 23. $\Gamma : \mathbf{Con}$ is a presheaf over \mathbb{C} . Its components are as follows.

$$\begin{aligned} |\Gamma| & : |\mathbb{C}| \rightarrow \mathbf{Set} \\ -\langle - \rangle & : |\Gamma| j \rightarrow \mathbb{C}(i, j) \rightarrow |\Gamma| i \\ \gamma\langle \mathbf{id} \rangle & = \gamma \\ \gamma\langle f \circ g \rangle & = \gamma\langle f \rangle \langle g \rangle \end{aligned}$$

We flip around the order of arguments in the action of Γ on morphisms. This is more convenient because of the contravariance; we can observe this in the statement of preservation laws already. The action on morphisms is sometimes called *restriction*.

Definition 24. $\sigma : \mathbf{Sub} \Gamma \Delta$ is a natural transformation from Γ to Δ . It has action $|\sigma| : |\Gamma| i \rightarrow |\Delta| i$, such that $|\sigma|(\gamma\langle f \rangle) = (|\sigma|\gamma)$.

Definition 25. $A : \mathbf{Ty} \Gamma$ is a displayed presheaf over Γ . The “displayed” here is used in exactly the same sense as in “displayed algebra” before. As we will see in Chapter 4, presheaves can be specified with a signature, in which case a presheaf is an algebra, and a displayed presheaf is a displayed algebra. The definition here is equivalent to saying that A is a presheaf over the category of elements of Γ ,

but it is more convenient to use in concrete definitions and calculations. The components of A are as follows.

$$\begin{aligned} |A| & : |\Gamma| i \rightarrow \mathbf{Set} \\ -\langle - \rangle & : |A| \gamma \rightarrow (f : \mathbb{C}(i, j)) \rightarrow |A| (\gamma \langle f \rangle) \\ \alpha \langle \mathbf{id} \rangle & = \alpha \\ \alpha \langle f \circ g \rangle & = \alpha \langle f \rangle \langle g \rangle \end{aligned}$$

Definition 26. $t : \mathbf{Tm} \Gamma A$ is a section of the displayed presheaf A . This is again the same notion of section that we have seen before, instantiated for presheaves.

$$\begin{aligned} |t| & : (\gamma : |\Gamma|) i \rightarrow |A| \gamma \\ |t|(\gamma \langle f \rangle) & = (|t| \gamma) \langle f \rangle \end{aligned}$$

Definition 27. $\Gamma \triangleright A : \mathbf{Con}$ is the total presheaf of the displayed presheaf A . Its action on objects and morphisms is the following.

$$\begin{aligned} |\Gamma \triangleright A| & \equiv (\gamma : |\Gamma|) \times |A| \gamma \\ (\gamma, \alpha) \langle f \rangle & \equiv (\gamma \langle f \rangle, \alpha \langle f \rangle) \end{aligned}$$

The \mathbf{id} and $-\circ-$ preservation laws follow immediately.

Definition 28. $A[\sigma] : \mathbf{Ty} \Gamma$ is defined as follows, assuming $A : \mathbf{Ty} \Delta$ and $\sigma : \mathbf{Sub} \Gamma \Delta$.

$$\begin{aligned} |A[\sigma]| \gamma & \equiv |A| (|\sigma| \gamma) \\ \alpha \langle f \rangle & \equiv \alpha \langle f \rangle \end{aligned}$$

In the second component, we use $-\langle - \rangle$ for A on the right hand side. The definition is well-typed since $|A| (|\sigma| (\gamma \langle f \rangle)) = |A| ((|\sigma| \gamma) \langle f \rangle)$ by the naturality of σ . Functoriality follows from functoriality of A .

It's easy to check that the above definitions can be extended to a cwf.

- For the base category, we take the category of presheaves.
- The empty context \bullet is the terminal presheaf, i.e. the constantly \top functor.
- Type substitution is functorial, as it is defined as simple function composition of actions on objects.

- Term substitution is defined as composition of a section and a natural transformation; and also functorial for the same reason.
- Context comprehension structure follows from the Σ -based definition for context extension.

Yoneda embedding

Before continuing with interpreting type formers in $\hat{\mathbb{C}}$, we review the Yoneda embedding, as it is useful in subsequent definitions.

Definition 29. The **Yoneda embedding**, denoted y , is a functor from \mathbb{C} to the underlying category of $\hat{\mathbb{C}}$, defined as follows.

$$\begin{aligned} y : |\mathbb{C}| &\rightarrow \mathbf{Con} & y : \mathbb{C}(i, j) &\rightarrow \mathbf{Sub}(y i)(y j) \\ y i &\equiv \mathbb{C}(-, i) & |y f| g &\equiv f \circ g \end{aligned}$$

Lemma 1 (Yoneda lemma). We have $\mathbf{Sub}(y i) \Gamma \simeq |\Gamma| i$ as an isomorphism of sets, natural in i [ML98, Section III.2].

Corollary. If we choose Γ to be $y j$, it follows that $\mathbf{Sub}(y i)(y j) \simeq \mathbb{C}(i, j)$, i.e. that y is bijective on morphisms; hence it is an embedding.

Notation 9. For $\gamma : |\Gamma| i$, we use $\gamma \langle - \rangle : \mathbf{Sub}(y i) \Gamma$ to denote transporting right-to-left along the Yoneda lemma. In the other direction we don't really need a notation, since from $\sigma : \mathbf{Sub}(y i) \Gamma$ we get $\sigma \text{id} : |\Gamma| i$.

Type formers

Definition 30. Constant families are displayed presheaves which do not depend on their context.

$$\begin{aligned} K & : \mathbf{Con} \rightarrow \{\Gamma : \mathbf{Con}\} \rightarrow \mathbf{Ty} \Gamma \\ |K \Delta| \{i\} \gamma & \equiv |\Delta| i \\ \delta \langle f \rangle & \equiv \delta \langle f \rangle \end{aligned}$$

With this definition, we have $\mathbf{Tm} \Gamma (K \Delta) = \mathbf{Sub} \Gamma \Delta$ so we have strict constant families.

Notation 10. It is useful to consider any set as a constant presheaf, so given $A : \mathbf{Set}$ we may write $A : \mathbf{Con}$ for the constant presheaf as well.

Definition 31. From any $A : \mathbf{Set}$, we get $\mathbf{K} A : \mathbf{Ty} \Gamma$. This can be used to model negative or positive **closed type formers**. For example, natural numbers are modeled as $\mathbf{K} \mathbb{N}$, Booleans as $\mathbf{K} \mathbf{Bool}$, the unit type as $\mathbf{K} \top$, and so on.

Definition 32. Coquand universes can be defined as follows. We write $\mathbf{Set}_{\hat{\mathbb{C}}}$ for the outer universe in the model, to distinguish it from the external \mathbf{Set} . Since the $\mathbf{Set}_{\hat{\mathbb{C}}}$ is a non-dependent type, it is helpful to define it as a $\mathbf{Set}_{\hat{\mathbb{C}}} : \mathbf{Con}$ such that $\mathbf{Sub} \Gamma \mathbf{Set}_{\hat{\mathbb{C}}} \simeq \mathbf{Ty} \Gamma$. The usual universe can be derived from this as $\mathbf{K} \mathbf{Set}_{\hat{\mathbb{C}}}$. Again, we ignore size issues; the fully formal definition would involve indexing constructions in $\hat{\mathbb{C}}$ by universe levels.

We can take a hint from the Yoneda lemma. We aim to define $|\mathbf{Set}_{\hat{\mathbb{C}}}| i$, but by the Yoneda lemma it is isomorphic to $\mathbf{Sub} (y i) \mathbf{Set}_{\hat{\mathbb{C}}}$. However, by specification this should be isomorphic to $\mathbf{Ty} (y i)$, so we take this as definition:

$$\begin{aligned} \mathbf{Set}_{\hat{\mathbb{C}}} & : \mathbf{Con} \\ |\mathbf{Set}_{\hat{\mathbb{C}}}| i & \equiv \mathbf{Ty} (y i) \\ A \langle f \rangle & \equiv A[y f] \end{aligned}$$

In the $A \langle f \rangle$ definition, we substitute $A : \mathbf{Ty} (y i)$ with $y f : \mathbf{Sub} (y j) (y i)$ to get an element of $\mathbf{Ty} (y j)$. The required $\mathbf{Sub} \Gamma \mathbf{Set}_{\hat{\mathbb{C}}} \simeq \mathbf{Ty} \Gamma$ is straightforward, so we omit the definition.

We note that *Russell* universes are not supported in the outer layer, as $\mathbf{Sub} \Gamma \mathbf{Set}_{\hat{\mathbb{C}}}$ and $\mathbf{Ty} \Gamma$ are not strictly the same, in particular they have a different number of components as iterated Σ -types. Nevertheless, as we mentioned in Section 3.3.2, we use Russell-style notation in the internal 2LTT syntax, and assume that encoding/decoding is inserted by elaboration.

Definition 33. Σ -types are defined pointwise. The definitions for pairing and projections follow straightforwardly.

$$\begin{aligned} \Sigma & : (A : \mathbf{Ty} \Gamma) \rightarrow \mathbf{Ty} (\Gamma \triangleright A) \rightarrow \mathbf{Ty} \Gamma \\ |\Sigma A B| \gamma & \equiv (\alpha : |A| \gamma) \times |B| (\gamma, \alpha) \\ (\alpha, \beta) \langle f \rangle & \equiv (\alpha \langle f \rangle, \beta \langle f \rangle) \end{aligned}$$

Definition 34. We define Π -types in the following. This is a bit more complicated, so first we look at the simpler case of presheaf exponentials. We source this example from [MM12, Section I.]. The reader may refer to *ibid.* for an overview of constructions in presheaf categories.

The exponential $\Delta^\Gamma : \mathbf{Con}$ is characterized by the isomorphism $\mathbf{Sub}(\Gamma \otimes \Delta) \Xi \simeq \mathbf{Sub} \Gamma(\Xi^\Delta)$, where we write \otimes for the pointwise product of two presheaves. We can again use the Yoneda lemma. We want to define $|\Delta^\Gamma| i$, but this is isomorphic to $\mathbf{Sub}(yi)(\Delta^\Gamma)$, which should be isomorphic to $\mathbf{Sub}(yi \otimes \Gamma) \Delta$ by the specification of exponentials. Hence:

$$\begin{aligned} |\Delta^\Gamma| i &\equiv \mathbf{Sub}(yi \otimes \Gamma) \Delta \\ \sigma\langle f \rangle &\equiv \sigma \circ (yf \circ \mathbf{p}, \mathbf{q}) \end{aligned}$$

In the definition of presheaf restriction, we use \mathbf{p}, \mathbf{q} as projections and $-, -$ as pairing for \otimes . In short, $(yf \circ \mathbf{p}, \mathbf{q})$ is the same as the morphism lifting from Definition 13: it weakens $yf : \mathbf{Sub}(yj)(yi)$ to $\mathbf{Sub}(yj \otimes \Gamma)(yi \otimes \Gamma)$.

The dependently typed case follows the same pattern, except that we use \mathbf{Tm} and $-\triangleright-$ instead of \mathbf{Sub} and $-\otimes-$. Additionally, the action on objects depends on $\gamma : |\Gamma| i$, and we make use of $\gamma\langle - \rangle : \mathbf{Sub}(yi) \Gamma$ (introduced in Notation 9).

$$\begin{aligned} \Pi & : (A : \mathbf{Tm} \Gamma) \rightarrow \mathbf{Tm}(\Gamma \triangleright A) \rightarrow \mathbf{Tm} \Gamma \\ |\Pi A B| \{i\} \gamma &\equiv \mathbf{Tm}(yi \triangleright A[\gamma\langle - \rangle])(B[\gamma\langle - \rangle \circ \mathbf{p}, \mathbf{q}]) \\ t\langle f \rangle & : \equiv t[yf \circ \mathbf{p}, \mathbf{q}] \end{aligned}$$

Let's unfold the above definition a bit. Assuming $t : |\Pi A B| \{i\} \gamma$, we have

$$|t| : \{j : |\mathbb{C}|\} \rightarrow ((f, \alpha) : (f : \mathbb{C}(j, i)) \times |A|(\gamma\langle f \rangle)) \rightarrow |B|(\gamma\langle f \rangle, \alpha)$$

This is a bit clearer if we remove the Σ -type by currying.

$$|t| : \{j : |\mathbb{C}|\}(f : \mathbb{C}(j, i))(\alpha : |A|(\gamma\langle f \rangle)) \rightarrow |B|(\gamma\langle f \rangle, \alpha)$$

Restriction is functorial since it is defined as \mathbf{Tm} substitution. The definitions for **lam** and **app** are left to the reader.

Definition 35. Extensional identity is defined as pointwise equality of sections:

$$\begin{aligned} \text{Id} & : \mathbf{Tm} \Gamma A \rightarrow \mathbf{Tm} \Gamma A \rightarrow \mathbf{Tm} \Gamma \\ |\text{Id } t u| \gamma &\equiv |t| \gamma = |u| \gamma \end{aligned}$$

For the restriction operation, we have to show that $|t| \gamma = |u| \gamma$ implies $|t|(\gamma\langle f \rangle) = |u|(\gamma\langle f \rangle)$. This follows from congruence by $\langle f \rangle$ and naturality of t and u . The defining $(\text{reflect}, \text{refl}) : \mathbf{Tm} \Gamma (\text{Id } t u) \simeq (t = u)$ isomorphism is evident, assuming UIP and function extensionality for the metatheoretic $=$ relation (which we do assume).

3.4.2 Modeling The Inner Layer

We assume now that \mathbb{C} is a cwf. We write types as $a, b, c : \text{Ty}_{\mathbb{C}} i$ and terms as $t, u, v : \text{Tm}_{\mathbb{C}} i a$. We reuse \bullet for the terminal object and $- \triangleright -$ for context extension, and likewise reuse notation for substitutions.

Definition 36 (Ty_0, Tm_0). First, note that $\text{Ty}_{\mathbb{C}}$ is a presheaf over \mathbb{C} , and $\text{Tm}_{\mathbb{C}}$ is a displayed presheaf over $\text{Ty}_{\mathbb{C}}$; this follows from the requirement that they form a family structure over \mathbb{C} . Hence, in the presheaf model $\text{Ty}_{\mathbb{C}}$ is an element of \mathbf{Con} and $\text{Tm}_{\mathbb{C}}$ is an element of $\text{Ty } \text{Ty}_{\mathbb{C}}$. Also recall from Definition 30 that $\text{Tm } \Gamma (\mathbf{K} \Delta) = \text{Sub } \Gamma \Delta$. With this in mind, we give the following definitions:

$$\begin{aligned} \text{Ty}_0 &: \text{Ty } \Gamma & \text{Tm}_0 &: \text{Tm } \Gamma \text{Ty}_0 \rightarrow \text{Ty } \Gamma \\ \text{Ty}_0 &:\equiv \mathbf{K } \text{Ty}_{\mathbb{C}} & \text{Tm}_0 A &:\equiv \text{Tm}_{\mathbb{C}}[A] \end{aligned}$$

$\text{Tm}_{\mathbb{C}}[A]$ is well-typed since $A : \text{Tm } \Gamma (\mathbf{K } \text{Ty}_{\mathbb{C}})$, thus $A : \text{Sub } \Gamma \text{Ty}_{\mathbb{C}}$. In other words, A is a natural transformation from Γ to the presheaf of inner types.

Inner type formers

Can type formers in $(\text{Ty}_{\mathbb{C}}, \text{Tm}_{\mathbb{C}})$ be transferred to $(\text{Ty}_0, \text{Tm}_0)$ in the presheaf model of 2LTT? For example, if \mathbb{C} supports **Bool**, we would like to model \mathbf{Bool}_0 in Ty_0 as well. The following explanation is adapted from Capriotti [Cap17, Section 2.3].

Generally, a type former in \mathbb{C} transfers to $\hat{\mathbb{C}}$ if it can be specified in the internal language of $\hat{\mathbb{C}}$; if the type former “always has been” in $\hat{\mathbb{C}}$ to begin with. To be describable in $\hat{\mathbb{C}}$, a type former needs to be natural with respect to \mathbb{C} morphisms. This is also a core idea of HOAS: when working in $\hat{\mathbb{C}}$, everything is natural, and we can omit boilerplate related to contexts and substitutions. For example, consider the specification of inner Π -types in the internal syntax of 2LTT:

$$\begin{aligned} \Pi_0 &: (A : \text{Ty}_0) \rightarrow (\text{Tm}_0 A \rightarrow \text{Ty}_0) \rightarrow \text{Ty}_0 \\ (\text{app}_0, \text{lam}_0) &: \Pi_0 A B \simeq ((a : \text{Tm}_0 A) \rightarrow \text{Tm}_0 (B a)) \end{aligned}$$

We can say that this *defines* what it means for \mathbb{C} to support Π . We recover the usual non-higher-order specification of Π in the following way, up to isomorphism:

- First, we interpret the higher-order specification as a context or closed Σ -type in the standard presheaf model of 2LTT. This yields a presheaf over \mathbb{C} .

- Then, we evaluate the resulting presheaf at the terminal object \bullet . This yields a set which is isomorphic to the conventional specification of Π .

In summary, if by “type formers” we mean extra structure on $(\mathsf{Ty}_0, \mathsf{Tm}_0)$ which is definable in 2LTT, then *by definition* all such type formers transfer from \mathbb{C} to $(\mathsf{Ty}_0, \mathsf{Tm}_0)$. This holds for every type former mentioned in this thesis.

3.4.3 Functions With Inner Domains

There is a useful semantic simplification in the standard presheaf model, in cases where we have functions of the form $\Pi (\mathsf{Tm}_0 A) B$. This greatly reduces encoding overhead when interpreting inductive signatures in 2LTT; we look at examples in Section 3.5. First we look at the simply-typed case with presheaf exponentials.

Lemma 2. y preserves finite products up to isomorphism, i.e. $y\bullet \simeq \bullet$ and $y(i \otimes j) \simeq (yi \otimes yj)$.

Proof. $y\bullet$ is $\mathbb{C}(-, \bullet)$ by definition, which is pointwise isomorphic to \top , hence isomorphic to $\bullet \equiv \mathsf{K} \top$. $y(i \otimes j)$ is $\mathbb{C}(-, i \otimes j)$, which is isomorphic to $yi \otimes yj$ by the specification of products. \square

Lemma 3. We have the following isomorphism.

$$\begin{aligned}
 |\Gamma^{yi}| j &\equiv \\
 \text{Sub } (yj \otimes yi) \Gamma &\simeq && \text{by product preservation} \\
 \text{Sub } (y(j \otimes i)) \Gamma &\simeq && \text{by Yoneda lemma} \\
 |\Gamma| (j \otimes i)
 \end{aligned}$$

It is possible to rephrase the above derivation for Π -types. For that, we would need to define the action of y on types and terms, consider the preservation of $- \triangleright -$ by y , and also specify a “dependent” Yoneda lemma for Tm . For the sake of brevity, we omit this, and present the result directly:

$$|\Pi (\mathsf{Tm}_0 A) B| \{i\} \gamma \simeq |B| \{i \triangleright |A| \gamma\} (\gamma \langle p \rangle, q)$$

In short, depending on an inner domain is the same as depending on an extended context in \mathbb{C} . We expand a bit on the typing of the right hand side. We have $\gamma : |\Gamma| i$, moreover

$$|B| : \{j : \mathbb{C}\} \rightarrow |\Gamma \triangleright \mathsf{Tm}_0 A| j \rightarrow \mathsf{Set}$$

$$\begin{aligned}
|B| & : \{j : \mathbb{C}\} \rightarrow ((\gamma' : |\Gamma| j) \times \mathbf{Tm}_{\mathbb{C}} j (|A| \gamma')) \rightarrow \mathbf{Set} \\
|B| \{i \triangleright |A| \gamma\} & : ((\gamma' : |\Gamma| (i \triangleright |A| \gamma)) \times \mathbf{Tm}_{\mathbb{C}} (i \triangleright |A| \gamma) (|A| \gamma')) \rightarrow \mathbf{Set} \\
\gamma \langle p \rangle & : |\Gamma| (i \triangleright |A| \gamma) \\
q & : \mathbf{Tm}_{\mathbb{C}} (i \triangleright |A| \gamma) ((|A| \gamma) [p]) \\
q & : \mathbf{Tm}_{\mathbb{C}} (i \triangleright |A| \gamma) (|A| (\gamma \langle p \rangle))
\end{aligned}$$

3.5 Simple Inductive Signatures in 2LTT

We revisit simple inductive signatures in this section, working internally to 2LTT. We review the concepts introduced in Chapter 2 in the same order.

Notation 11. In this section we shall be fairly explicit about writing \mathbf{Tm}_0 -s and transporting along definitional isomorphisms. The simple setting makes it feasible to be explicit; in later chapters we are more terse, as signatures and semantics get more complicated.

3.5.1 Theory of Signatures

Signatures are defined exactly in the same way as before: we have $\mathbf{Con} : \mathbf{Set}, \mathbf{Ty} : \mathbf{Set}, \mathbf{Sub} : \mathbf{Con} \rightarrow \mathbf{Con} \rightarrow \mathbf{Set}, \mathbf{Var} : \mathbf{Con} \rightarrow \mathbf{Ty} \rightarrow \mathbf{Set}$ and $\mathbf{Tm} : \mathbf{Con} \rightarrow \mathbf{Ty} \rightarrow \mathbf{Set}$. However, now by \mathbf{Set} we mean the outer universe of 2LTT. Thus signatures are inductively defined in the outer layer.

3.5.2 Algebras

Again we compute algebras by induction on signatures, but now we use inner types for carriers of algebras. We interpret types as follows:

$$\begin{aligned}
-^A & : \mathbf{Ty} \rightarrow \mathbf{Ty}_0 \rightarrow \mathbf{Set} \\
\iota^A & \quad X : \equiv \mathbf{Tm}_0 X \\
(\iota \rightarrow A)^A & X : \equiv \mathbf{Tm}_0 X \rightarrow A^A X
\end{aligned}$$

Elsewhere, we change the type of the X parameters accordingly:

$$\begin{aligned} -^A &: \text{Con} \rightarrow \text{Ty}_0 \rightarrow \text{Set} \\ -^A &: \text{Var } \Gamma \ A \rightarrow \{X : \text{Ty}_0\} \rightarrow \Gamma^A X \rightarrow A^A X \\ -^A &: \text{Tm } \Gamma \ A \rightarrow \{X : \text{Ty}_0\} \rightarrow \Gamma^A X \rightarrow A^A X \\ -^A &: \text{Sub } \Gamma \ \Delta \rightarrow \{X : \text{Ty}_0\} \rightarrow \Gamma^A X \rightarrow \Delta^A X \end{aligned}$$

We also define $\text{Alg } \Gamma$ as $(X : \text{Ty}_0) \times \Gamma^A X$.

Example 8. Inside 2LTT we have the following:³

$$\text{Alg NatSig} \equiv (X : \text{Ty}_0) \times (\text{zero} : \text{Tm}_0 X) \times (\text{suc} : \text{Tm}_0 X \rightarrow \text{Tm}_0 X)$$

Then, we may assume any cwf \mathbb{C} , and interpret the above closed type in the presheaf model $\hat{\mathbb{C}}$, and evaluate the result at \bullet and the unique element of the terminal presheaf $\mathbf{K} \top$:

$$|\text{Alg NatSig}| \{\bullet\} \text{tt} : \text{Set}$$

We only need to compute definitions now. We use the simplified semantics for $\text{suc} : \text{Tm}_0 X \rightarrow \text{Tm}_0 X$, since the function domain is an inner type. We get the following:

$$|\text{Alg NatSig}| \{\bullet\} \text{tt} \equiv (X : \text{Ty}_{\mathbb{C}} \bullet) \times (\text{zero} : \text{Tm}_{\mathbb{C}} \bullet X) \times (\text{suc} : \text{Tm}_{\mathbb{C}} (\bullet \triangleright X) X)$$

Using the same computation, we get the following for binary trees:

$$|\text{Alg TreeSig}| \{\bullet\} \text{tt} \equiv (X : \text{Ty}_{\mathbb{C}} \bullet) \times (\text{leaf} : \text{Tm}_{\mathbb{C}} \bullet X) \times (\text{node} : \text{Tm}_{\mathbb{C}} (\bullet \triangleright X \triangleright X) X)$$

We can also get internal algebras in any \mathbb{C} category with finite products, because we can build cwfs from all such \mathbb{C} .

Definition 37. Assuming \mathbb{C} with finite products, we build a cwf by setting $\text{Con} \equiv |\mathbb{C}|$, $\text{Ty } \Gamma \equiv |\mathbb{C}|$, $\text{Sub } \Gamma \ \Delta \equiv \mathbb{C}(\Gamma, \Delta)$, $\text{Tm } \Gamma \ A \equiv \mathbb{C}(\Gamma, A)$, $\Gamma \triangleright A \equiv \Gamma \otimes A$ and $\bullet \equiv \bullet_{\mathbb{C}}$. In short, we build a non-dependent (simply-typed) cwf.

Now we can effectively interpret signatures in finite product categories. For example:

$$|\text{Alg NatSig}| \{\bullet\} \text{tt} \equiv (X : |\mathbb{C}|) \times (\text{zero} : \mathbb{C}(\bullet, X)) \times (\text{suc} : \mathbb{C}(\bullet \otimes X, X))$$

This is almost the same as what we would write by hand for the specification of natural number objects; the only difference is the extra $\bullet \otimes -$ in suc .

³Up to isomorphism, since we previously defined Γ^A as a function type instead of an iterated product type.

3.5.3 Morphisms

We get an additional degree of freedom in the computation of morphisms: preservation equations can be inner or outer. The former option is *weak* or *propositional* preservation, while the latter is *strict* preservation. In the presheaf model of 2LTT, outer equality is definitional equality of inner terms, while inner equality is propositional equality in the inner theory. Of course, if the inner theory has extensional identity types, weak and strict equations in 2LTT coincide for inner types. We compute weak preservation for types as follows.

$$\begin{aligned} -^M : (A : \mathbf{Ty})\{X_0 X_1 : \mathbf{Ty}_0\}(X^M : \mathbf{Tm}_0 X_0 \rightarrow \mathbf{Tm}_0 X_1) &\rightarrow A^A X_0 \rightarrow A^A X_1 \rightarrow \mathbf{Set} \\ \iota^M X^M \alpha_0 \alpha_1 &\equiv \mathbf{Tm}_0 (X^M \alpha_0 = \alpha_1) \\ (\iota \rightarrow A)^M X^M \alpha_0 \alpha_1 &\equiv (x : \mathbf{Tm}_0 X_0) \rightarrow A^M X^M (\alpha_0 x) (\alpha_1 (X^M x)) \end{aligned}$$

For strict preservation, we simply change $\mathbf{Tm}_0 (X^M \alpha_0 = \alpha_1)$ to $X^M \alpha_0 \equiv \alpha_1$. The definition of morphisms is the same as before:

$$\begin{aligned} -^M : (\Gamma : \mathbf{Con}_1)\{X_0 X_1 : \mathbf{Ty}_0\} &\rightarrow (\mathbf{Tm}_0 X_0 \rightarrow \mathbf{Tm}_0 X_1) \rightarrow \Gamma^A X_0 \rightarrow \Gamma^A X_1 \rightarrow \mathbf{Set} \\ \Gamma^M X^M \gamma_0 \gamma_1 &\equiv \{A\}(x : \mathbf{Var}_1 \Gamma A) \rightarrow A^M X^M (\gamma_0 x) (\gamma_1 x) \end{aligned}$$

$$\mathbf{Mor} : \{\Gamma : \mathbf{Con}_1\} \rightarrow \mathbf{Alg} \Gamma \rightarrow \mathbf{Alg} \Gamma \rightarrow \mathbf{Set}$$

$$\mathbf{Mor} \{\Gamma\} (X_0, \gamma_0) (X_1, \gamma_1) \equiv (X^M : \mathbf{Tm}_0 X_0 \rightarrow \mathbf{Tm}_0 X_1) \times \Gamma^M X^M \gamma_0 \gamma_1$$

We omit here the $-^M$ definitions for terms and substitutions.

3.5.4 Displayed Algebras

We present $-^D$ only for types below.

$$\begin{aligned} -^D : (A : \mathbf{Ty})\{X\} &\rightarrow (\mathbf{Tm}_0 X \rightarrow \mathbf{Ty}_0) \rightarrow A^A X \rightarrow \mathbf{Set} \\ \iota^D X^D \alpha &\equiv \mathbf{Tm}_0 (X^D \alpha) \\ (\iota \rightarrow A)^D X^D \alpha &\equiv (x : \mathbf{Tm}_0 X)(x^D : \mathbf{Tm}_0 (X^D x)) \rightarrow A^D X^D (\alpha x) \end{aligned}$$

We use type dependency in the inner layer. In the presheaf model, this is interpreted as inner types depending on certain contexts.

Example 9. Assume a closed $(X, \text{zero}, \text{suc})$ Nat-algebra in 2LTT. We have the

following computation:

$$\begin{aligned} \text{DispAlg } \{\text{NatSig}\} (X, \text{zero}, \text{suc}) &\equiv \\ (X^D &: \mathsf{Tm}_0 X \rightarrow \mathsf{Ty}_0) \\ \times (\text{zero}^D &: \mathsf{Tm}_0 (X^D \text{zero})) \\ \times (\text{suc}^D &: (n : \mathsf{Tm}_0 X) \rightarrow \mathsf{Tm}_0 (X^D n) \rightarrow \mathsf{Tm}_0 (X^D (\text{suc } n))) \end{aligned}$$

Let's look at the presheaf interpretation now. We simplify functions with inner domains everywhere. Also note that for $\text{suc} : \mathsf{Tm}_0 X \rightarrow \mathsf{Tm}_0 X$, we get $|\text{suc}| \mathsf{tt} : \mathsf{Tm}_{\mathbb{C}} (\bullet \triangleright n : |X| \mathsf{tt}) (|X| \mathsf{tt})$ in the semantics, so a $\text{suc } t$ application is translated as a substitution $(|\text{suc}| \mathsf{tt})[n \mapsto |t| \mathsf{tt}]$.

$$\begin{aligned} &|\text{DispAlg } \{\text{NatSig}\} (X, \text{zero}, \text{suc})| \{\bullet\} \mathsf{tt} \equiv \\ &(X^D : \mathsf{Ty}_{\mathbb{C}} (\bullet \triangleright n : |X| \mathsf{tt})) \\ &\times (\text{zero}^D : \mathsf{Tm}_{\mathbb{C}} \bullet (X^D [n \mapsto |\text{zero}| \mathsf{tt}])) \\ &\times (\text{suc}^D : \mathsf{Tm}_{\mathbb{C}} (\bullet \triangleright n : |X| \mathsf{tt} \triangleright n^D : X^D [n \mapsto |\text{zero}| \mathsf{tt}]) (X^D [n \mapsto (|\text{suc}| \mathsf{tt})[n \mapsto n]])) \end{aligned}$$

To explain $(|\text{suc}| \mathsf{tt})[n \mapsto n]$: we have $\text{suc } n$ in 2LTT, where n is an inner variable, and in the presheaf model inner variables become actual variables in the inner theory. Hence, we map the n which suc depends on to the concrete n in the context.

We can also interpret displayed algebras in finite product categories:

$$\begin{aligned} &|\text{DispAlg } \{\text{NatSig}\} (X, \text{zero}, \text{suc})| \{\bullet\} \mathsf{tt} \equiv \\ &(X^D : |\mathbb{C}|) \\ &\times (\text{zero}^D : \mathbb{C}(\bullet, X^D)) \\ &\times (\text{suc}^D : \mathbb{C}(\bullet \otimes |X| \mathsf{tt} \otimes X^D, X^D)) \end{aligned}$$

While displayed algebras in cwfs can be used as bundles of induction motives and methods, in finite product categories they are argument bundles to *primitive recursion*; this is sometimes also called a *paramorphism* [MFP91]. In an internal syntax, the type of primitive recursion for natural numbers could be written more compactly as:

$$\text{primrec} : (X : \mathsf{Set}) \rightarrow X \rightarrow (\mathsf{Nat} \rightarrow X \rightarrow X) \rightarrow \mathsf{Nat} \rightarrow X$$

This is not the same thing as the usual recursion principle (corresponding to weak initiality), because of the extra dependency on Nat in the method for successors.

3.5.5 Sections

Sections are analogous to morphisms. We again have a choice between weak and strict preservation; below we have weak preservation.

$$\begin{aligned}
& -^S : (A : \mathbf{Ty}) \{X X^D\} (X^S : (x : \mathbf{Tm}_0 X) \rightarrow \mathbf{Tm}_0 (X^D x)) \\
& \rightarrow (\alpha : A^A X) \rightarrow A^D X^D \alpha \rightarrow \mathbf{Set} \\
& \iota^S \quad X^S \alpha \alpha^D :\equiv \mathbf{Tm}_0 (X^S \alpha = \alpha^D) \\
& (\iota \rightarrow A)^S X^S \alpha \alpha^D :\equiv (x : \mathbf{Tm}_0 X) \rightarrow A^S X^S (\alpha x) (\alpha^D (X^S x))
\end{aligned}$$

3.5.6 Term Algebras

For term algebras, we need to assume a bit more in the inner theory. For starters, it has to support the theory of signatures. In order to avoid name clashes down the line, we use \mathbf{SigTy}_0 to refer to signature types, and \mathbf{SigTm}_0 for terms. That is, we have

$$\begin{aligned}
& \mathbf{SigTy}_0 : \mathbf{Ty}_0 \\
& \mathbf{Con}_0 : \mathbf{Ty}_0 \\
& \mathbf{Var}_0 : \mathbf{Tm}_0 \mathbf{Con}_0 \rightarrow \mathbf{Tm}_0 \mathbf{SigTy}_0 \rightarrow \mathbf{Ty}_0 \\
& \mathbf{SigTm}_0 : \mathbf{Tm}_0 \mathbf{Con}_0 \rightarrow \mathbf{Tm}_0 \mathbf{SigTy}_0 \rightarrow \mathbf{Ty}_0 \\
& \mathbf{Sub}_0 : \mathbf{Tm}_0 \mathbf{Con}_0 \rightarrow \mathbf{Tm}_0 \mathbf{Con}_0 \rightarrow \mathbf{Ty}_0
\end{aligned}$$

together with all constructors and induction principles. We also assume inner Π -types, because we previously defined \mathbf{Sub} using functions.

Remark. If we only want to construct term algebras, it is not necessary to assume inner induction principles. It is sometimes useful to construct term algebras from models which are explicitly *not* initial; we will see such examples in ???. In this section, our goal is to redo the constructions of Chapter 2 without making essential changes, so we just assume everything that was available there.

We still have ToS in the outer layer. To make the naming scheme consistent, we shall write outer ToS types as \mathbf{SigTy}_1 , \mathbf{SigTm}_1 , \mathbf{Con}_1 , \mathbf{Var}_1 and \mathbf{Sub}_1 . We have conversion functions from the outer ToS to the inner ToS:

Definition 38. We have the following **lowering** functions which preserve all

structure.

$$\begin{aligned}
\downarrow : \text{SigTy}_1 &\rightarrow \text{Tm}_0 \text{SigTy}_0 \\
\downarrow : \text{Con}_1 &\rightarrow \text{Tm}_0 \text{Con}_0 \\
\downarrow : \text{Var}_1 \Gamma A &\rightarrow \text{Tm}_0 (\text{Var}_0 (\downarrow \Gamma) (\downarrow A)) \\
\downarrow : \text{SigTm}_1 \Gamma A &\rightarrow \text{Tm}_0 (\text{SigTm}_0 (\downarrow \Gamma) (\downarrow A)) \\
\downarrow : \text{Sub}_1 \Gamma \Delta &\rightarrow \text{Tm}_0 (\text{Sub}_0 (\downarrow \Gamma) (\downarrow \delta))
\end{aligned}$$

These functions are called “lifting” in the context of multi-stage programming; see e.g. the `Lift` typeclass in Haskell [PWK19]. There, like here, the point is to build object-language terms from meta-level (“compile-time”) values.

Lowering is straightforward to define for types, contexts, variables and terms, but there is a bit of a complication for `Sub`. Unfolding the definitions, we need to map from $\{A\} \rightarrow \text{Var}_1 \Delta A \rightarrow \text{SigTm}_1 \Gamma A$ to $\text{Tm}_0 (\{A\} \rightarrow \text{Var}_0 (\downarrow \Delta) A \rightarrow \text{SigTm}_0 (\downarrow \Gamma) A)$. It might appear problematic that we have types and variables in *negative* position, because we can’t map inner types/variables to outer ones. Fortunately, $\text{Sub}_1 \Gamma \Delta$ is isomorphic to a finite product type, and we can lower a finite product component-wise.

Concretely, we define lowering by induction on Δ , while making use of a case splitting operation for Var_0 . We use an informal `case` operation below, which can be defined using inner induction. Note that since $\text{Var}_0 \bullet A$ is an empty type, case splitting on it behaves like elimination for the empty type.

$$\begin{aligned}
\downarrow_\Delta : \text{Sub}_1 \Gamma \Delta &\rightarrow \text{Tm}_0 (\text{Sub}_0 (\downarrow \Gamma) (\downarrow \Delta)) \\
\downarrow_\bullet \sigma &\equiv \lambda \{A\} (x : \text{Var}_0 \bullet A). \text{case } x \text{ of } () \\
\downarrow_{\Delta \triangleright B} \sigma &\equiv \lambda \{A\} (x : \text{Var}_1 (\downarrow \Delta \triangleright \downarrow B) A). \text{case } x \text{ of} \\
\text{vz} &\rightarrow \downarrow (\sigma \text{vz}) \\
\text{vs } x &\rightarrow \downarrow_\Delta (\sigma \circ \text{vs}) x
\end{aligned}$$

In general, for finite A type, functions of the form $A \rightarrow \text{Tm}_0 B$ can be represented as inner types up to isomorphism; they can be viewed as finite products of terms.

Remark. For infinite A this does not work anymore in our system. In [ACKS19], the assumption that this still works with $A \equiv \text{Nat}_1$ is an important axiom (“cofibrancy of Nat_1 ”) which makes it possible to embed higher categorical structures in 2LTT. From the metaprogramming perspective, cofibrancy of Nat_1 implies that the inner theory is *infinitary*, since we can form inner terms from

infinite collections of inner terms. We don't assume this axiom in 2LTT, although we will consider infinitary (object) type theories in Chapters 4 and 5.

We continue to the definition of term algebras. We fix an $\Omega : \mathbf{Con}_1$, and define $\mathsf{T} : \mathsf{Ty}_0$ as $\mathsf{SigTm}_0(\downarrow\Omega)\iota$.

$$\begin{aligned} -^T &: (A : \mathsf{SigTy}_1) \rightarrow \mathsf{Tm}_0(\mathsf{SigTm}_0(\downarrow\Omega)(\downarrow A)) \rightarrow A^A \mathsf{T} \\ \iota^T &\quad t \equiv t \\ (\iota \rightarrow A)^T t &\equiv \lambda u. A^T(\mathsf{app} \, t \, u) \end{aligned}$$

$$\begin{aligned} -^T &: (\Gamma : \mathbf{Con}_1) \rightarrow \mathbf{Sub}_1 \Omega \Gamma \rightarrow \Gamma^A \mathsf{T} \\ \Gamma^T \nu \{A\} x &\equiv A^T(\downarrow(\nu x)) \end{aligned}$$

$$\begin{aligned} \mathsf{TmAlg}_\Omega &: \mathbf{Alg} \, \Omega \\ \mathsf{TmAlg}_\Omega &\equiv \Omega^T \Omega \mathsf{id} \end{aligned}$$

We omit the $-^T$ interpretation for terms and substitutions for now, as they require a bit more setup, and they are not needed just for term algebras.

3.5.7 Recursor Construction

Recall from Section 2.3.1 that recursion is implemented using the $-^A$ interpretation of terms. Since terms are now in the inner theory, we need to define an inner version of the same interpretation. We need to compute types by inner induction, so we additionally assume a Russell-style inner U_0 universe. The Russell style means that we may freely coerce between $\mathsf{Tm}_0 \mathsf{U}_0$ and Ty_0 . The following are defined the same way as $-^A$ before.

$$\begin{aligned} -^A &: \mathsf{Tm}_0(\mathsf{SigTy}_0 \rightarrow \mathsf{U}_0 \rightarrow \mathsf{U}_0) \\ -^A &: \mathsf{Tm}_0(\mathbf{Con}_0 \rightarrow \mathsf{U}_0 \rightarrow \mathsf{U}_0) \\ -^A &: \mathsf{Tm}_0(\mathsf{SigTm}_0 \Gamma A \rightarrow \{X : \mathsf{U}_0\} \rightarrow \Gamma^A X \rightarrow A^A X) \\ -^A &: \mathsf{Tm}_0(\mathbf{Sub}_0 \Gamma \Delta \rightarrow \{X : \mathsf{U}_0\} \rightarrow \Gamma^A X \rightarrow \Delta^A X) \end{aligned}$$

Since lowering preserves all structure, and $-^A$ is defined in the same way in both the inner and outer theories, lowering is compatible with $-^A$ in the following way.

Lemma 4. Assume $A : \mathsf{SigTy}_1$, $\Gamma : \mathbf{Con}_1$, $X : \mathsf{Ty}_0$, $\gamma : \Gamma^A X$ and $t : \mathsf{SigTm}_1 \Gamma A$. We have the following:

- $(A_{\rightarrow}^A, A_{\leftarrow}^A) : \mathbf{Tm}_0((\downarrow A)^A X) \simeq A^A X$
- $(\Gamma_{\rightarrow}^A, \Gamma_{\leftarrow}^A) : \mathbf{Tm}_0((\downarrow \Gamma)^A X) \simeq \Gamma^A X$
- $t^A \gamma \equiv A_{\rightarrow}^A((\downarrow t)^A(\Gamma_{\leftarrow}^A \gamma))$

Proof. By induction on Γ , A and t . □

We construct recursors now, yielding strict algebra morphisms. We assume $(X, \omega) : \mathbf{Alg} \Omega$. Recall that $\omega : \Omega^A X$, thus $\Omega_{\leftarrow}^A \omega : \mathbf{Tm}_0((\downarrow \Omega)^A X)$. We define $R : \mathbf{Tm}_0 \mathbf{T} \rightarrow \mathbf{Tm}_0 X$ as $R t \equiv t^A(\Omega_{\leftarrow}^A \omega)$.

$$\begin{aligned} -^R : (A : \mathbf{SigTy}_1)(t : \mathbf{Tm}_0(\mathbf{SigTm}_0(\downarrow \Omega)(\downarrow A))) &\rightarrow A^M R(A^T t)(A_{\rightarrow}^A(t^A(\Omega_{\leftarrow}^A \omega))) \\ \iota_{\rightarrow}^R \quad t : t^A(\Omega_{\leftarrow}^A \omega) &\equiv \iota_{\rightarrow}^A(t^A(\Omega_{\leftarrow}^A \omega)) \\ (\iota \rightarrow A)^R t &\equiv \lambda u. A^R(\mathbf{app} t u) \end{aligned}$$

$$\begin{aligned} -^R : (\Gamma : \mathbf{Con}_1)(\nu : \mathbf{Sub}_1 \Omega \Gamma) &\rightarrow \Gamma^M R(\Gamma^T \nu)(\nu^A \omega) \\ \Gamma^R \nu \{A\} x &\equiv A^R(\downarrow(\nu x)) \end{aligned}$$

In the proof obligation for $t^A(\Omega_{\leftarrow}^A \omega) \equiv \iota_{\rightarrow}^A(t^A(\Omega_{\leftarrow}^A \omega))$, ι_{\rightarrow}^A computes to the identity function; note that $\iota_{\rightarrow}^A : \mathbf{Tm}_0 X \rightarrow \mathbf{Tm}_0 X$. Hence the equality becomes reflexive.

In $\Gamma^R \nu \{A\} x \equiv A^R(\downarrow(\nu x))$, we have that

$$A^R \downarrow(\nu x) : A^M R(A^T(\downarrow(\nu x)))(A_{\rightarrow}^A(\downarrow(\nu x)^A(\Omega_{\leftarrow}^A \omega)))$$

Hence by Lemma 4, we have

$$A^R \downarrow(\nu x) : A^M R(A^T(\downarrow(\nu x)))((\nu x)^A \omega)$$

Hence, by the definition of $-^A$ for substitutions:

$$A^R \downarrow(\nu x) : A^M R(A^T(\downarrow(\nu x))) (\nu^A \omega x)$$

Which is exactly what is required when we unfold the expected return type:

$$\begin{aligned} -^R : (\Gamma : \mathbf{Con}_1)(\nu : \mathbf{Sub}_1 \Omega \Gamma) &\rightarrow \Gamma^M R(\Gamma^T \nu)(\nu^A \omega) \\ -^R : (\Gamma : \mathbf{Con}_1)(\nu : \mathbf{Sub}_1 \Omega \Gamma) &\rightarrow \{A\}(x : \mathbf{Var}_1 \Gamma A) \rightarrow A^M R(A^T(\downarrow(\nu x))) (\nu^A \omega x) \end{aligned}$$

The recursor is defined the same way as in Definition 10:

$$\begin{aligned} \mathbf{Rec}_{\Omega} : (alg : \mathbf{Alg} \Omega) &\rightarrow \mathbf{Mor} \mathbf{TmAlg}_{\Omega} alg \\ \mathbf{Rec}_{\Omega}(X, \omega) &\equiv (R, \Omega^R \Omega \text{id}) \end{aligned}$$

3.5.8 Eliminator Construction

For induction, we need to additionally define $-^D$ in the inner layer.

$$\begin{aligned} -^D &: \mathbf{Tm}_0((A : \mathbf{SigTy}_0)\{X\} \rightarrow (\mathbf{Tm}_0 X \rightarrow \mathbf{U}_0) \rightarrow A^A X \rightarrow \mathbf{U}_0) \\ -^D &: \mathbf{Tm}_0((\Gamma : \mathbf{Con}_0)\{X\} \rightarrow (\mathbf{Tm}_0 X \rightarrow \mathbf{U}_0) \rightarrow \Gamma^A X \rightarrow \mathbf{U}_0) \\ -^D &: \mathbf{Tm}_0((t : \mathbf{SigTm}_0 \Gamma A) \rightarrow \Gamma^D X^D \gamma \rightarrow A^D X^D (t^A \gamma)) \\ -^D &: \mathbf{Tm}_0((\sigma : \mathbf{Sub}_0 \Gamma \Delta) \rightarrow \Gamma^D X^D \gamma \rightarrow \Delta^D X^D (\sigma^A \gamma)) \end{aligned}$$

Lemma 5. We have again compatibility of lowering with $-^D$. Assuming $(X, \gamma) : \mathbf{Alg} \Gamma$, $(X^D, \gamma^D) : \mathbf{DispAlg} (X, \gamma)$, $t : \mathbf{SigTm}_1 \Gamma A$, and $\alpha : A^A X$, we have

- $(A_{\rightarrow}^D, A_{\leftarrow}^D) : \mathbf{Tm}_0((\downarrow A)^D X^D (A_{\leftarrow}^A \alpha)) \simeq A^D X^D \alpha$
- $(\Gamma_{\rightarrow}^D, \Gamma_{\leftarrow}^D) : \mathbf{Tm}_0((\downarrow \Gamma)^D X^D (\Gamma_{\leftarrow}^A \gamma)) \simeq \Gamma^D X^D \gamma$
- $t^D \gamma^D \equiv A_{\rightarrow}^D((\downarrow t)^D (\Gamma_{\leftarrow}^D \gamma^D))$

The equation for $t^D \gamma^D$ is well-typed because of the term equation in Lemma 4.

Proof. Again by induction on Γ , A and t . \square

We also need to extend $-^T$ with action on terms. Note that we return an inner equality, since we can only compute such equality by induction on the inner term input:

$$-^T : (t : \mathbf{SigTm}_0(\downarrow \Gamma)(\downarrow A))(\nu : \mathbf{Sub}_1 \Omega \Gamma) \rightarrow \mathbf{Tm}_0(A_{\leftarrow}^A(A^T(t[\downarrow \nu]))) = t^A(\Gamma_{\leftarrow}^A \nu))$$

We assume $(X^D, \omega^D) : \mathbf{DispAlg} \mathbf{TmAlg}_\Omega$, and define elimination as follows:

$$\begin{aligned} \mathbf{E} &: (t : \mathbf{Tm}_0 \mathbf{T}) \rightarrow \mathbf{Tm}_0(X^D t) \\ \mathbf{E} t &\equiv t^D(\Omega_{\leftarrow}^D \omega^D) \end{aligned}$$

This definition is well-typed only up to $t^T \mathbf{id} : \mathbf{Tm}_0(t = t^A(\Omega_{\leftarrow}^A(\Omega^T \Omega \mathbf{id})))$. Since $t^T \mathbf{id}$ is an inner equality, in a fully formal intensional presentation we would have to write an explicit transport in the definition.

We shall skip the remainder of the eliminator construction; it goes the same way as in Definition 11. Intuitively, this is possible since the inner theory has all necessary features to reproduce the eliminator construction, and lowering preserves all structure.

Since t^T yields inner equations, this implies that the displayed algebra sections returned by the eliminator are *weak sections*, i.e. they contain β -rules expressed in inner equalities.

3.6 Discussion

3.6.1 Evaluation

Let's review how 2LTT addresses the shortcomings that we mentioned in Section 3.1.

Generalized semantics. For this purpose, there is minimal technical overhead to using 2LTT; we only need to sprinkle Ty_0 and Tm_0 around a bit, and we get semantics internally to cwfs (including all finite product categories).

Term algebra constructions. Here, we had to deal with significantly more noise, because of the necessary lowerings and their preservation isomorphisms. However, this overhead should be compared to reasoning about object-level definitional equality in a deeply embedded way, which entails explicitly manipulating abstract syntax and its substitutions and weakenings. Compared to that, 2LTT is tremendously easier. One could also argue that lowering is an entirely mechanical affair, and we can just omit most of it if we are comfortable enough with the formalism.

3.6.2 Recursor vs. Eliminator Construction

The essential extra detail that we had to handle in Section 3.5 was the choice between strict and weak equations. This choice brings along further implementation constraints.

Strict equations are stronger as assumptions, because they represent definitional equality of inner terms. However, we can only produce strict equations by eliminating from outer types. Hence, if we aim to output strict equations, we have to assume every dependency in the outer theory, which in turn may require using lowering.

Weak equations are easier to produce: strict equality implies weak equality, plus we can prove weak equality by inner induction. But we cannot use weak equality to transport outer values.

In the recursor construction, we produced strict β -rules, which trivially imply weak β -rules as well. In contrast, the eliminator construction relies essentially on equations produced by $-^T$ for inner terms. Can we somehow get strict equations

from $-^T$? This does not seem possible, at least without changing the approach in a major way: terms must be inner, or else we have no hope of inner recursion/induction. And if terms are inner, then \mathbf{E} must act on arbitrary inner terms, hence $-^T$ must do so too.

It would be interesting to check if there is a possible alternative formulation of term algebras and constructed eliminators, which makes it possible to get strict eliminators. Looking back to Section 2.4.2, it appears that by using spine neutral terms, we get strict elimination for each signature. However, this relies on the Agda implementation of pattern matching and structurally recursive definitions, so it would require more work to translate these definitions to 2LTT in a generic way.

CHAPTER 4

Finitary Quotient Inductive-Inductive Signatures

In this chapter we bump the expressive power of signatures by a large margin, and also substantially extend the semantics. However, we keep the basic approach the same; indeed its advantages become apparent with the more sophisticated signatures.

We use two different setups for semantics in this chapter.

- In Sections 4.1-4.2 we work in 2LTT, thereby getting a generalized semantics for signatures. However, we keep details about universe levels to the minimum.
- In Sections ??-4.4, we work in an extensional type theory with cumulative universes. This is more suited for the term algebra construction, where (as we will see) 2LTT does not bring any advantage, but we do need to be more precise about universes.

4.1 Theory of Signatures

Signatures are once again given by contexts of a type theory, but now it is a dependent type theory, given as a cwf with certain type formers, in the style of Section 3.2.

Metatheory and notation. We work in 2LTT with Ty_0 and Tm_0 . We make the following assumptions. Later, we may assume more type formers at either levels as needed.

- Ty_0 is closed under \top , Σ and extensional identity $- = -$. The inner identity reflects the outer one.
- The outer identity $- \equiv -$ is also extensional; it supports UIP, function extensionality, and it reflects strict equality in some unspecified metatheory outside 2LTT. This reflection is used to justify omitting transports along $- \equiv -$ in our notation.

In the following we specify models of the theory of *finitary quotient inductive-inductive signatures*. The names involved are a bit of a mouthful, so we abbreviate “finitary quotient inductive-inductive” as FQII, and like before, we abbreviate “theory of signatures” as ToS. In this chapter, by signature we mean an FQII signature unless otherwise specified.

Definition 39. A **model of the theory of signatures** consists of the following.

- A **cwf** with underlying sets Con , Sub , Ty and Tm , all returning in the outer Set universe of 2LTT.
- A **Tarski-style universe** U with decoding El .
- An **extensional identity type** $\mathsf{Id} : \mathsf{Tm} \Gamma A \rightarrow \mathsf{Tm} \Gamma A \rightarrow \mathsf{Ty} \Gamma$, specified by $(\mathsf{reflect}, \mathsf{refl}) : \mathsf{Tm} \Gamma (\mathsf{Id} \, t \, u) \simeq (t \equiv u)$.
- An **inductive function type** $\Pi : (a : \mathsf{Tm} \Gamma \mathsf{U}) \rightarrow \mathsf{Ty} (\Gamma \triangleright \mathsf{El} \, a) \rightarrow \mathsf{Ty} \Gamma$, specified by $(\mathsf{app}, \mathsf{lam}) : \mathsf{Tm} \Gamma (\Pi \, a \, B) \simeq \mathsf{Tm} (\Gamma \triangleright \mathsf{El} \, a) \, B$.
- An **external function type** $\Pi^{\mathsf{ext}} : (A : \mathsf{Ty}_0) \rightarrow (A \rightarrow \mathsf{Ty} \Gamma) \rightarrow \mathsf{Ty} \Gamma$, specified by $(\mathsf{app}^{\mathsf{ext}}, \mathsf{lam}^{\mathsf{ext}}) : \mathsf{Tm} \Gamma (\Pi^{\mathsf{ext}} \, A \, B) \simeq ((x : A) \rightarrow \mathsf{Tm} \Gamma (B \, x))$.

At this point we only have a notion of model for ToS, but as we will see in Chapter 5, ToS is also an algebraic theory, more specifically an infinitary QII one. It is infinitary because Π^{ext} and $\mathsf{lam}^{\mathsf{ext}}$ allow branching over elements of arbitrary $A : \mathsf{Ty}_0$ types.

Because of the algebraic character of ToS, there is a category of ToS models where morphisms strictly preserve all structure, and the initial model corresponds to the syntax. In this chapter, we just assume that the ToS syntax exists, and leave metatheoretical matters to Chapter 5.

Definition 40. An **FQII signature** is an element of Con in the syntax of ToS.

We review several example signatures in the following, using progressively more ToS type formers. We also introduce progressively more compact notation for signatures. As a rule of thumb, we shall use compact notation for larger and more complex signatures, but we shall be more explicit when we specify models of ToS later in this chapter.

Example 10. Simple inductive signatures can be evidently expressed using \mathbf{U} and Π . By adding a single \mathbf{U} to the signature, we introduce the inductive sort.

$$\mathbf{NatSig} : \equiv \bullet \triangleright (N : \mathbf{U}) \triangleright (\mathit{zero} : \mathbf{El} N) \triangleright (\mathit{suc} : \Pi(n : N)(\mathbf{El} N))$$

$$\mathbf{TreeSig} : \equiv \bullet \triangleright (T : \mathbf{U}) \triangleright (\mathit{leaf} : \mathbf{El} T) \triangleright (\mathit{node} : \Pi(t_1 : T)(\Pi(t_2 : T)(\mathbf{El} T)))$$

Observe that the domains in Π are terms with type \mathbf{U} , while the codomains are proper types.

Notation 12. We write non-dependent function types in ToS as follows.

- $a \Rightarrow B$ for $\Pi(- : a) B$.
- $A \Rightarrow^{\text{ext}} B$ for $\Pi^{\text{ext}} A (\lambda _ . B)$.

Using this notation, we may write $\mathit{suc} : N \Rightarrow \mathbf{El} N$ and $\mathit{node} : T \Rightarrow T \Rightarrow \mathbf{El} T$.

Notation 13. The “categorical” application \mathbf{app} with explicit substitutions is a bit inconvenient. Instead, we simply write whitespace for Π and Π^{ext} application:

$$t u : \equiv (\mathbf{app} t)[\mathit{id}, u]$$

$$t u : \equiv (\mathbf{app}^{\text{ext}} t) u$$

Example 11. We may have any number of sorts by adding more \mathbf{U} to the signatures. Moreover, sorts can be indexed over previous sorts. Hence, using only \mathbf{U} , \mathbf{El} and Π , we can express any closed inductive-inductive type [?]. The following fragment of the the signature for categories is such:

$$\bullet \triangleright (\mathit{Obj} : \mathbf{U}) \triangleright (\mathit{Hom} : \mathit{Obj} \Rightarrow \mathit{Obj} \Rightarrow \mathbf{U}) \triangleright (\mathit{id} : \Pi(i : \mathit{Obj})(\mathbf{El}(\mathit{Hom} i i)))n$$

These inductive-inductive signatures are more flexible than those in prior works [?], since we allow type constructors (sorts) and point constructors to be arbitrarily mixed (as opposed to mandating that sorts are declared first). For example:

$$\bullet \triangleright (A : \mathbf{U}) \triangleright (a : \mathbf{El} A) \triangleright (B : A \Rightarrow \mathbf{U}) \triangleright (C : B a \Rightarrow \mathbf{U})$$

Here C is indexed over $B a$, where a is a point constructor of a , so a sort specification mentions a point constructor.

Example 12. `ld` lets us add equations to signatures. With this, we can write down the full signature for categories:

$$\begin{aligned}
&\bullet \triangleright (\text{Obj} : \mathbf{U}) \\
&\triangleright (\text{Hom} : \text{Obj} \Rightarrow \text{Obj} \Rightarrow \mathbf{U}) \\
&\triangleright (\text{id} : \Pi(i : \text{Obj}) (\mathbf{El} (\text{Hom } i \ i))) \\
&\triangleright (\text{comp} : \Pi(i \ j \ k : \text{Obj}) (\text{Hom } i \ j \Rightarrow \text{Hom } i \ k \Rightarrow \mathbf{El} (\text{Hom } i \ k))) \\
&\triangleright (\text{idr} : \Pi(i \ j : \text{Obj}) (f : \text{Hom } i \ j) (\mathbf{ld} (\text{comp } i \ i \ j \ f (\text{id } i)) \ f)) \\
&\triangleright (\text{idl} : \Pi(i \ j : \text{Obj}) (f : \text{Hom } i \ j) (\mathbf{ld} (\text{comp } i \ j \ j (\text{id } j) \ f) \ f)) \\
&\triangleright (\text{assoc} : \Pi(i \ j \ k \ l : \text{Obj}) (f : \text{Hom } j \ l) (g : \text{Hom } j \ k) (h : \text{Hom } i \ j) \\
&\quad (\mathbf{ld} (\text{comp } i \ j \ l (\text{comp } j \ k \ l \ f \ g) \ h) (\text{comp } i \ k \ l \ f (\text{comp } i \ j \ k \ g \ h))))
\end{aligned}$$

Now, this is already rather hard to read, even together with the compressed notation for multiple Π binders.

Notation 14. For more complex signatures, we may entirely switch to an internal notation, where we mostly reuse the conventions in the metatheories. We use $(x : a) \rightarrow B$ for inductive functions, $(x : A) \rightarrow^{\text{ext}} B$ for external functions, but we still write `ld` for the identity type and make \mathbf{U} and \mathbf{El} explicit. In this notation, a signature is just a listing of binders. The category signature becomes the following:

$$\begin{aligned}
&\text{Obj} : \mathbf{U} \\
&\text{Hom} : \text{Obj} \rightarrow \text{Obj} \rightarrow \mathbf{U} \\
&\text{id} : \mathbf{El} (\text{Hom } i \ i) \\
&- \circ - : \text{Hom } j \ k \rightarrow \text{Hom } i \ j \rightarrow \mathbf{El} (\text{Hom } i \ k) \\
&\text{idr} : \mathbf{ld} (f \circ \text{id}) \ f \\
&\text{idl} : \mathbf{ld} (\text{id} \circ f) \ f \\
&\text{assoc} : \mathbf{ld} (f \circ (g \circ h)) ((f \circ g) \circ h)
\end{aligned}$$

Example 13. The external function type makes it possible to reference inner types (in 2LTT) in signatures. Here “external” is meant relative to a given signature, and refers to types and inhabitants which are not introduced inside a signature. For example, we give a signature for lists by assuming $A : \mathbf{Ty}_0$ for the

(external) type of list elements:

$$\begin{aligned} List & : \mathbf{U} \\ nil & : \mathbf{El} \, List \\ cons & : A \rightarrow^{\text{ext}} List \rightarrow \mathbf{El} \, List \end{aligned}$$

Hence, “parameters” are always assumptions made in the metatheory. We can also *index* sorts by external values. Let’s specify length-indexed vectors now; we keep the $A : \mathbf{Ty}_0$ assumption, but also assume that \mathbf{Ty}_0 has natural numbers, with $\mathbf{Nat}_0 : \mathbf{Ty}_0$, \mathbf{zero}_0 and \mathbf{suc}_0 .

$$\begin{aligned} Vec & : \mathbf{Nat}_0 \rightarrow^{\text{ext}} \mathbf{U} \\ nil & : \mathbf{El} \, (Vec \, \mathbf{zero}_0) \\ cons & : (n : \mathbf{Nat}_0) \rightarrow^{\text{ext}} A \rightarrow^{\text{ext}} Vec \, n \rightarrow \mathbf{El} \, (Vec \, (\mathbf{suc}_0 \, n)) \end{aligned}$$

Example 14. We can also introduce *sort equations* using **ld**: this means equating terms of \mathbf{U} , i.e. inductively specified sets. This is useful for specifying certain strict type formers. For example, a signature for cwfs can be extended with a specification for strict constant families.

$$\begin{aligned} Con & : \mathbf{U} \\ Sub & : Con \rightarrow Con \rightarrow \mathbf{U} \\ Ty & : Con \rightarrow \mathbf{U} \\ Tm & : (\Gamma : Con) \rightarrow Ty \, \Gamma \rightarrow \mathbf{U} \\ \dots & \\ K & : Con \rightarrow \{\Gamma : Con\} \rightarrow \mathbf{El} \, (Ty \, \Gamma) \\ K_{spec} & : \mathbf{ld} \, (Tm \, \Gamma \, (K \, \Delta)) \, (Sub \, \Gamma \, \Delta) \end{aligned}$$

The equation for Russell-style universes is likewise a sort equation:

$$\begin{aligned} Univ & : \mathbf{El} \, (Ty \, \Gamma) \\ Russell & : \mathbf{ld} \, (Tm \, \Gamma \, Univ) \, (Ty \, \Gamma) \end{aligned}$$

Example 15. As we mentioned in Definition 25, there is a signature for presheaves, so let’s look at that now. Assume a category \mathbb{C} in the inner theory; this means

that objects and morphisms of \mathbb{C} are in \mathbf{Ty}_0 .

$$\begin{aligned} Obj & : |\mathbb{C}| \rightarrow^{\text{ext}} \mathbf{U} \\ Hom & : \mathbb{C}(i, j) \rightarrow^{\text{ext}} Obj\ j \rightarrow \mathbf{El}(Obj\ i) \\ Hom_{\text{id}} & : \mathbf{Id}(Hom\ \text{id}\ x)\ x \\ Hom_{\circ} & : \mathbf{Id}(Hom\ (f \circ g)\ x)\ (Hom\ f\ (Hom\ g\ x)) \end{aligned}$$

We depart from the sugary naming scheme in Definition 25, and name the action on objects Obj and the action on morphisms Hom . When we give semantics to this signature in Section 4.2, we will get as algebras functors from \mathbb{C}^{op} to the category of inner types. That category has elements of \mathbf{Ty}_0 as objects and $\mathbf{Tm}_0 A \rightarrow \mathbf{Tm}_0 B$ functions as morphisms.

Strict positivity. Only strictly positive signatures are expressible. Similarly to the case with simple signatures, there is no way to abstract over a inductive functions, since inductive function domains are in \mathbf{U} , and \mathbf{U} has no type formers at all. With Π^{ext} , we can abstract over functions, but only on ones which are external to a signature and don't depend on internally specified constructors.

4.2 Semantics

4.2.1 Overview

For simple signatures, we only gave semantics in enough detail so that notions of recursion and induction could be recovered. We aim to do more now. For each signature, we would like to have

1. A *category* of algebras, with homomorphisms as morphisms.
2. A notion of induction, which requires a notion of dependent algebras.
3. A proof that for algebras, initiality is equivalent to supporting induction.

We do this by creating a model of ToS where contexts (signatures) are categories with certain extra structure and substitutions are structure-preserving functors. Then, ToS signatures can be interpreted in this model, using the initiality of ToS syntax (i.e. the recursor).

Our semantics has a type-theoretic flavor, which is inspired by the cubical set model of Martin-Löf type theory by Bezem et al. [?]. The idea is to avoid strictness

issues by starting from basic ingredients which are already strict enough. Hence, instead of modeling ToS types as certain slices and substitution by pullback, we model types as displayed categories with extra structure, which naturally support strict reindexing/substitution.

We make a similar choice in the interpretation of signatures themselves: we use structured cwfs of algebras, where types correspond to displayed algebras. This choice is in contrast to having finitely complete categories of algebras. Preliminarily, the reason is that “native” displayed algebras and sections allow us to compute induction principles strictly as one would write in a type theory. In fact, in this chapter we recover exactly the same semantics for simple signatures that we already specified.

In contrast, displayed algebras are a derived notion in finitely complete categories, and the induction principles would be only up to isomorphism. This issue is perhaps not relevant from a purely categorical perspective, but we are concerned with eventually implementing QIITs in proof assistants. If we don’t compute induction principles here in an exact way, we don’t get them from anywhere else.

4.2.2 Separate vs. Bundled Models

Previously, we defined $-^A$, $-^M$, $-^D$ and $-^S$ interpretations of signatures separately, by doing induction anew for each one. Formally, this amounts to giving a plain model of ToS in order to define $-^A$, but then giving three *displayed* models of ToS to specify the other interpretations, because they sometimes need to refer to the recursors or eliminators of other interpretations.

For example, $-^A : \mathbf{Con} \rightarrow \mathbf{Set}$ while $-^D : (\Gamma : \mathbf{Con}) \rightarrow \Gamma^A \rightarrow \mathbf{Set}$, so displayed algebras already refer to $-^A$, which is part of the recursor for the corresponding model.

However, this piecewise style can be avoided: we can give a single non-displayed model which packs everything in a Σ -type, yielding just one interpretation function for signatures. Let’s call that function $-^M$ now:

$$\begin{aligned} -^M : \mathbf{Con} \rightarrow & (A : \mathbf{Set}) \\ & \times (M : A \rightarrow A \rightarrow \mathbf{Set}) \\ & \times (D : A \rightarrow \mathbf{Set}) \\ & \times (S : (a : A) \rightarrow D a \rightarrow \mathbf{Set}) \end{aligned}$$

Note that it is often not possible to merge multiple recursors/eliminators by packing models together. For example, addition on natural numbers is defined by recursion, and so is multiplication; but since multiplication calls addition in an iterated fashion, it is not possible to define both operations by a single algebra. Nevertheless, merging does work in our case. We will, in fact, get a formal vocabulary for merging models (and manipulating them in other ways) from the semantics of ToS itself.

In simple cases, and in Agda, the piecewise style is convenient, since we don't have to deal with Σ -s. However, for larger models, important organizing principles may become more apparent if we bundle things together.

In the following, we shall define a model $\mathbf{M} : \text{ToS}$ such that its **Con** component is a bundle containing all A, M, D, S components, plus a number of additional components. We present the components of \mathbf{M} in the same order as in Definition 39. There is significant overlap in names and notations, so we use **bold** font to disambiguate components of \mathbf{M} from components of other structures. For example, we use $\sigma : \mathbf{Sub} \Gamma \Delta$ to denote a substitution in \mathbf{M} , while there could be a **Sub**-named components in other structures under consideration.

4.2.3 Finite Limit Cwfs

We define **Con** : **Set** as the type of finite limit cwfs. Recall that this specifies the objects of the underlying cwf of \mathbf{M} . In the following we specify flcwfs and describe some internal constructions.

Definition 41. We define **flcwf** : **Set** as an iterated Σ -type with the following components:

1. A cwf with **Con**, **Sub**, **Ty**, **Tm** all returning in **Set**. *Remark:* this implies that **flcwf** : **Set** is in a larger universe than all of these internal components. We continue to elide universe sizing details.
2. Σ -types.
3. Extensional identity type **ld** with **refl** and **reflect**.
4. Strict constant families **K**.

Definition 42. We abbreviate the additional structure on cwfs consisting of Σ , **ld** and **K** as **fl-structure**.

We recover previous concepts as follows. Assuming Γ signature, we get an *flwf* by interpreting Γ in \mathbf{M} . In that *flwf* we have

- **Con** as the type of algebras.
- **Sub** as the type of algebra morphisms.
- **Ty** as the type of displayed algebras.
- **Tm** as the type of displayed algebra sections.

From this, notions of initiality and induction are apparent as well. Initiality is the usual categorical notion. For induction, assuming $\Gamma : \mathbf{Con}$, we have the following predicate:

$$\begin{aligned} \text{Inductive} &: \mathbf{Con}_\Gamma \rightarrow \mathbf{Set} \\ \text{Inductive } \Gamma &\equiv (A : \mathbf{Ty}_\Gamma \Gamma) \rightarrow \mathbf{Tm}_\Gamma \Gamma A \end{aligned}$$

In short, an algebra is inductive if every displayed algebra over it has a section. Fortunately, we also know that induction and initiality are equivalent.

Theorem 1. *An object $\Gamma : \mathbf{Con}_\Gamma$ in an *flwf* Γ supports induction if and only if it is initial. Moreover, induction and initiality are both mere properties.*

Proof. First, we show that induction implies initiality. We assume $\Gamma : \mathbf{Con}$, $\text{ind} : \text{Inductive } \Gamma$ and $\Delta : \mathbf{Con}$. We aim to show that there is a unique inhabitant of $\mathbf{Sub } \Gamma \Delta$. We have $\text{ind } (\mathbf{K } \Delta) : \mathbf{Tm } \Gamma (\mathbf{K } \Delta)$, hence $\text{ind } (\mathbf{K } \Delta) : \mathbf{Sub } \Gamma \Delta$. We only need to show that this is unique. Assume $\delta : \mathbf{Sub } \Gamma \Delta$. Now, $\text{ind } (\text{Id } \delta (\text{ind } (\mathbf{K } \Delta))) : \mathbf{Tm } \Gamma (\text{Id } \delta (\text{ind } (\mathbf{K } \Delta)))$, and it follows by equality reflection that $\delta \equiv \text{ind } (\mathbf{K } \Delta)$.

Second, the other direction. We assume that Γ is initial, and also $A : \mathbf{Ty } \Gamma$, and aim to inhabit $\mathbf{Tm } \Gamma A$. By initiality we get a unique $\sigma : \mathbf{Sub } \Gamma (\Gamma \triangleright A)$. Now, $\mathbf{q}[\sigma] : \mathbf{Tm } \Gamma (A[\mathbf{p} \circ \sigma])$, but since $\mathbf{p} \circ \sigma : \mathbf{Sub } \Gamma \Gamma$, it must be equal to id by the initiality of Γ . Hence, $\mathbf{q}[\sigma] : \mathbf{Tm } \Gamma A$.

Lastly: it is well-known that initiality is a mere property, so let's show the same for induction. We assume $\text{ind}, \text{ind}' : \text{Inductive } \Gamma$ and $A : \mathbf{Ty } \Gamma$. We have $\text{reflect } (\text{ind } (\text{Id } (\text{ind } A) (\text{ind}' A))) : \text{ind } A \equiv \text{ind}' A$. Since A is arbitrary, by function extensionality we also have $\text{ind} \equiv \text{ind}'$. \square

Theorem 2. *$\mathbf{Tm } \Gamma A$ in an *flwf* is propositional when Γ is initial.*

Proof. Assuming $t, u : \mathbf{Tm } \Gamma A$, we have $\text{reflect } (\text{ind } (\text{Id } t u)) : t \equiv u$. \square

Note that the above proof does not rely on Σ -types in the flcwf, so why do we include them in the semantics? One reason is the prior result by Clairmabault and Dybjer [CD14] that a slightly different formulation of flcwf is biequivalent to finitely complete categories. More concretely, in *ibid.* there is a 2-category of cwf with Σ , Id and “democracy”, the last of which is equivalent to the weak formulation of constant families. Then, it is shown that this 2-category is biequivalent to the 2-category of finitely complete categories. Thus, including Σ is a good deal, as this allows us to connect our semantics back to finitely complete categories, which are more commonplace in categorical settings.

We recover finite limits in an flcwf as follows. The product of Γ and Δ is given by $\Gamma \triangleright \mathbf{K} \Delta$, and we get projection and pairing from context comprehension. The equalizer of $\sigma, \delta : \mathbf{Sub} \Gamma \Delta$ is given by $\Gamma \triangleright \text{Id } \sigma \delta$, which is well-typed because morphisms can be viewed as terms, e.g. $\sigma : \mathbf{Tm} \Gamma (\mathbf{K} \Delta)$. The unique morphism out of the equalizer is $\mathbf{p} : \mathbf{Sub} (\Gamma \triangleright \text{Id } \sigma \delta) \Gamma$.

Our flcwf is not exactly the same as in [CD14] because our constant families are strict. But we certainly don’t lose anything by having stricter semantics, since the weak version can be trivially recovered.

In the following we present some concepts and results in flcwf.

Definition 43 (Type categories, c.f. [CD14, Section 2.2]). For each $\Gamma : \mathbf{Con}$, there is a category whose objects are types $A : \mathbf{Ty} \Gamma$, and morphisms from A to B are terms $t : \mathbf{Tm} (\Gamma \triangleright A) (B[\mathbf{p}])$. Identity morphisms are given by $\mathbf{q} : \mathbf{Tm} (\Gamma \triangleright A) (A[\mathbf{p}])$, and composition $t \circ u$ by $t[\mathbf{p}, u]$. The assignment of type categories to contexts extends to a split indexed category. For each $\sigma : \mathbf{Sub} \Gamma \Delta$, there is a functor from $\mathbf{Ty} \Delta$ to $\mathbf{Ty} \Gamma$, which sends A to $A[\sigma]$ and $t : \mathbf{Tm} (\Gamma \triangleright A) (B[\mathbf{p}])$ to $t[\sigma \circ \mathbf{p}, \mathbf{q}]$.

Notation 15.

- In any cwf, we use $\sigma : \Gamma \simeq \Delta$ to indicate that $\sigma : \mathbf{Sub} \Gamma \Delta$ is an isomorphism with inverse σ^{-1} .
- A *type isomorphism*, written as $t : A \simeq B$ is an isomorphism in a type category, with inverse as t^{-1} .

Theorem 3 (Equivalence of types and slices, c.f. [CD14, Section 2.2]). *Assume that we work in an flcwf Γ . For each $\Gamma : \mathbf{Con}$, the type category $\mathbf{Ty} \Gamma$ is equivalent to the slice category Γ/Γ .*

Remark. In the flw of sets where types are $A \rightarrow \mathbf{Set}$ families, the above theorem yields the equivalence of $A \rightarrow \mathbf{Set}$ and $(B : \mathbf{Set}) \times (B \rightarrow A)$. This is sometimes called the “family-fibration” equivalence. It is also a notable motivating example for univalence in type theory: it is not an isomorphism of sets, but only an equivalence up to isomorphism of sets. So this is an example for an equivalence which quite naturally arises even if we only care about sets, but one which is not covered by set-level univalence, and actually requires univalence for groupoids, if we want to prove it as a propositional equality.

4.2.4 The Cwf of Finite Limit Cwfs

The next task is to define the cwf part of \mathbf{M} . We already know that objects are flwfs.

Category

A **morphism** $\sigma : \mathbf{Sub} \Gamma \Delta$ is in an algebra homomorphism, viewing flwfs as algebraic structures. Hence, σ includes a functor between underlying categories, but it also maps types to types and terms to terms, and *strictly preserves all structure*.

Notation 16. We may implicitly project out the underlying maps from σ . Hence, we have the following four maps:

$$\begin{aligned} \sigma &: \mathbf{Con}_\Gamma \rightarrow \mathbf{Con}_\Delta \\ \sigma &: \mathbf{Sub}_\Gamma \Gamma \Delta \rightarrow \mathbf{Sub}_\Delta (\sigma \Gamma) (\sigma \Delta) \\ \sigma &: \mathbf{Ty}_\Gamma \Gamma \rightarrow \mathbf{Ty}_\Delta (\sigma \Gamma) \\ \sigma &: \mathbf{Tm}_\Gamma \Gamma A \rightarrow \mathbf{Tm}_\Delta (\sigma \Gamma) (\sigma A) \end{aligned}$$

We list some of the preservation equations as examples of usage:

$$\begin{aligned} \sigma \bullet &\equiv \bullet \\ \sigma (\Gamma \triangleright A) &\equiv \sigma \Gamma \triangleright \sigma A \\ \sigma (A[\sigma]) &\equiv (\sigma A)[\sigma \sigma] \\ \sigma (t[\sigma]) &\equiv (\sigma t)[\sigma \sigma] \\ \sigma (\Sigma A B) &\equiv \Sigma (\sigma A) (\sigma B) \\ \sigma (\text{proj}_1 t) &\equiv \text{proj}_1 (\sigma t) \end{aligned}$$

Above, we could have also included subscripts indicating the Γ or Δ flw, as in $\sigma \bullet_{\Gamma} \equiv \bullet_{\Delta}$; but these are quite easily inferable, so we omit them.

Identity morphisms and **composition** are defined in the evident way using identity functions and function composition in underlying maps, and they satisfy the category laws.

The terminal object $\bullet : \mathbf{Con}$ is given by having $\mathbf{Con}_{\bullet} := \top$, $\mathbf{Sub}_{\bullet} \Gamma \Delta := \top$, $\mathbf{Ty}_{\bullet} \Gamma := \top$ and $\mathbf{Tm}_{\bullet} \Gamma A := \top$, and all structure and equations are defined trivially.

Family structure

A type $A : \mathbf{Ty} \Gamma$ is a displayed flw over Γ . As we have seen before, displayed algebras can be computed as logical predicate interpretations of algebraic signatures. Every A component lies over the corresponding Γ component. Also note that a displayed flw includes a displayed category, for which some results have been worked out in [AL19].

Notation 17. In situations where we need to refer to both “base” and displayed things, we give underlined names to contexts, substitutions, types and terms in a base flw. For example, we may have $\underline{\Gamma} : \mathbf{Con}_{\Gamma}$ living in $\Gamma : \mathbf{Con}$, and $\Gamma : \mathbf{Con}_A \underline{\Gamma}$ living in a displayed flw over Γ . We only use underlining on 2LTT variable names, and overload flw component names for displayed counterparts. For example, a \mathbf{Con} component is named the same in a base flw and a displayed one.

Concretely, a displayed flw A over Γ has the following underlying sets, which we call displayed contexts, substitutions, types and terms respectively.

$$\begin{aligned} \mathbf{Con}_A &: \mathbf{Con}_{\Gamma} \rightarrow \mathbf{Set} \\ \mathbf{Sub}_A &: \mathbf{Con}_A \underline{\Gamma} \rightarrow \mathbf{Con}_A \underline{\Delta} \rightarrow \mathbf{Sub}_{\Gamma} \underline{\Gamma} \underline{\Delta} \rightarrow \mathbf{Set} \\ \mathbf{Ty}_A &: \mathbf{Con}_A \underline{\Gamma} \rightarrow \mathbf{Ty}_{\Gamma} \underline{\Gamma} \rightarrow \mathbf{Set} \\ \mathbf{Tm}_A &: (\Gamma : \mathbf{Con}_A \underline{\Gamma}) \rightarrow \mathbf{Ty}_A \underline{\Gamma} \underline{A} \rightarrow \mathbf{Tm}_{\Gamma} \underline{\Gamma} \underline{A} \rightarrow \mathbf{Set} \end{aligned}$$

We list select components of A below; note how every A operation lies over the corresponding Γ operation. In our notation with implicit arguments, equations in A can be written the same way as in Γ , but of course there is extra indexing involved, and the displayed equations are well-typed because of their counterparts

in the base.

$$\begin{aligned}
\text{id}_A & : \text{Sub } \Gamma \Gamma \text{id}_\Gamma \\
- \circ_A - & : \text{Sub } \Delta \Xi \underline{\sigma} \rightarrow \text{Sub } \Gamma \Delta \underline{\delta} \rightarrow \text{Sub } \Gamma \Xi (\underline{\sigma} \circ_\Gamma \underline{\delta}) \\
\text{idl}_A & : \text{id}_A \circ_A \sigma \equiv \sigma \\
\text{idr}_A & : \sigma \circ_A \text{id}_A \equiv \sigma \\
\bullet_A & : \text{Con}_A \bullet_\Gamma \\
- \triangleright_A - & : (\Gamma : \text{Con}_A \underline{\Gamma}) \rightarrow \text{Ty}_A \Gamma \underline{A} \rightarrow \text{Con}_A \Gamma (\underline{\Gamma} \triangleright_\Gamma \underline{A}) \\
-[-]_A & : \text{Ty}_A \Delta \underline{A} \rightarrow \text{Sub}_A \Gamma \Delta \underline{\sigma} \rightarrow \text{Ty}_A \Gamma (\underline{A}[\underline{\sigma}]_\Gamma) \\
-[-]_A & : \text{Tm}_A \Delta A \underline{t} \rightarrow (\sigma : \text{Sub}_A \Gamma \Delta \underline{\sigma}) \rightarrow \text{Tm}_A \Gamma (A[\sigma]_A) (\underline{t}[\underline{\sigma}]_\Gamma) \\
\text{Id}_A & : \text{Tm}_A \Gamma A \underline{t} \rightarrow \text{Tm}_A \Gamma A \underline{u} \rightarrow \text{Ty}_A \Gamma (\text{Id}_\Gamma \underline{t} \underline{u}) \\
K_A & : \text{Con}_A \underline{\Delta} \rightarrow \{\Gamma : \text{Con}_A \underline{\Gamma}\} \rightarrow \text{Ty}_A \Gamma (K_\Gamma \underline{\Delta}) \\
\Sigma_A & : (A : \text{Ty}_A \Gamma \underline{A}) \rightarrow \text{Ty}_A (\Gamma \triangleright_A A) \underline{B} \rightarrow \text{Ty}_A \Gamma (\Sigma_\Gamma \underline{A} \underline{B})
\end{aligned}$$

In the following we will often omit Γ and A subscripts on components; for example, in the type $\text{Con}_A \bullet$, the \bullet is clearly a base component in Γ .

A substituted type $A[\sigma] : \text{Ty } \Gamma$ is defined as follows, for $A : \text{Ty } \Delta$ and $\sigma : \text{Sub } \Gamma \Delta$. We simply compose underlying functions in σ with the underlying predicates in A :

$$\begin{aligned}
\text{Con}_{A[\sigma]} \underline{\Gamma} & : \equiv \text{Con}_A (\sigma \underline{\Gamma}) \\
\text{Sub}_{A[\sigma]} \Gamma \Delta \underline{\sigma} & : \equiv \text{Sub}_A \Gamma \Delta (\sigma \underline{\sigma}) \\
\text{Ty}_{A[\sigma]} \Gamma \underline{A} & : \equiv \text{Ty}_A \Gamma (\sigma \underline{A}) \\
\text{Tm}_{A[\sigma]} \Gamma A \underline{t} & : \equiv \text{Tm}_A \Gamma A (\sigma \underline{t})
\end{aligned}$$

It should be clear that $A[\sigma]$ thus defined still supports all displayed flowf structure. For example, the displayed contexts in $A[\sigma]$ are elements of $\text{Con}_A (\sigma \underline{\Gamma})$, but since σ preserves all Γ -structure, we can also recover all displayed structure. For example, if $\underline{\Gamma}$ is \bullet , we have $\sigma \bullet \equiv \bullet$, and we can reuse $\bullet_A : \text{Con}_A \bullet$ to define the displayed empty context in $A[\sigma]$, and we can proceed analogously for all other structure in $A[\sigma]$.

Additionally, type substitution is functorial, i.e. $A[\text{id}] \equiv A$ and $A[\sigma \circ \delta] \equiv A[\sigma][\delta]$. This holds because the underlying set families are defined by function composition.

Remark. Types could be equivalently defined as slices in the category of flows, and type substitution could be given as pullback, but in that case we would run

into the well-known strictness issue, that type substitution is functorial only up to isomorphism [?]. This is not a critical issue, as there are standard solutions for recovering strict substitutions from weak ones [?]. But if we ever need to look inside the definitions in the model, using displayed algebras yields a lot less encoding overhead than strictifying pullbacks.

A term $t : \mathbf{Tm} \Gamma \mathbf{A}$ is a displayed flwf section, which again strictly preserves all structure. We use the same notation for the action of t that we use for **Sub**. We have the following underlying maps:

$$\begin{aligned} t : (\underline{\Gamma} : \mathbf{Con}_{\Gamma}) &\rightarrow \mathbf{Con}_{\mathbf{A}} \underline{\Gamma} \\ t : (\underline{\sigma} : \mathbf{Sub}_{\Gamma} \Gamma \Delta) &\rightarrow \mathbf{Sub}_{\mathbf{A}} (t \Gamma) (t \Delta) \underline{\sigma} \\ t : (\underline{A} : \mathbf{Ty}_{\Gamma} \Gamma) &\rightarrow \mathbf{Ty}_{\mathbf{A}} (t \Gamma) \underline{A} \\ t : (\underline{t} : \mathbf{Tm}_{\Gamma} \Gamma A) &\rightarrow \mathbf{Tm}_{\mathbf{A}} (t \Gamma) (t A) \underline{t} \end{aligned}$$

A substituted term $t[\sigma]$ for $t : \mathbf{Tm} \Delta \mathbf{A}$ and $\sigma : \mathbf{Sub} \Gamma \Delta$ is again given by component-wise function composition.

An extended context $\Gamma \triangleright \mathbf{A}$ is the *total flwf* of \mathbf{A} . This is defined by combining corresponding underlying sets with Σ -types:

$$\begin{aligned} \mathbf{Con}_{\Gamma \triangleright \mathbf{A}} &:\equiv (\underline{\Gamma} : \mathbf{Con}_{\Gamma}) \times \mathbf{Con}_{\mathbf{A}} \underline{\Gamma} \\ \mathbf{Sub}_{\Gamma \triangleright \mathbf{A}} (\underline{\Gamma}, \Gamma) (\underline{\Delta}, \Delta) &:\equiv (\underline{\sigma} : \mathbf{Sub}_{\Gamma} \underline{\Gamma} \underline{\Delta}) \times \mathbf{Sub}_{\mathbf{A}} \Gamma \Delta \underline{\sigma} \\ \mathbf{Ty}_{\Gamma \triangleright \mathbf{A}} (\underline{\Gamma}, \Gamma) &:\equiv (\underline{A} : \mathbf{Ty}_{\Gamma} \underline{\Gamma}) \times \mathbf{Ty}_{\mathbf{A}} \Gamma \underline{A} \\ \mathbf{Tm}_{\Gamma \triangleright \mathbf{A}} (\underline{\Gamma}, \Gamma) (\underline{A}, A) &:\equiv (\underline{t} : \mathbf{Tm}_{\Gamma} \underline{\Gamma} \underline{A}) \times \mathbf{Tm}_{\mathbf{A}} \Gamma A \underline{t} \end{aligned}$$

All structure is defined pointwise, using Γ -structure for first projections and \mathbf{A} -structure for second projections. $\Gamma \triangleright \mathbf{A}$ may be viewed as a dependent generalization of direct products of flwfs.

Comprehension structure follows from the above definition: \mathbf{p} is component-wise first projection, \mathbf{q} is second projection and substitution extension $-, -$ is pairing.

With this, we have a cwf of flwfs. *Remark:* flwf itself is algebraic, and has a finitary QII signature. Hence, if we succeed building semantics for finitary QII signatures, we get “for free” an flwf of flwfs. Of course, we cannot rely on this when we’re in the process of defining the **M** model in the first place. Checking that the **M** model indeed works, is the somewhat tedious task that we have to perform *once*, in order to get semantics for any other finitary QII theory.

4.2.5 Type Formers

Strict constant families

This was not included in the ToS specification, but it's quite useful, so we shall define it. $\mathbf{K} \Delta : \mathbf{Ty} \Gamma$ is defined by ignoring Γ inhabitants in all underlying sets:

$$\begin{aligned} \text{Con}_{\mathbf{K} \Delta} \underline{\Gamma} &: \equiv \text{Con}_{\Delta} \\ \text{Sub}_{\mathbf{K} \Delta} \Gamma \Delta \underline{\sigma} &: \equiv \text{Sub}_{\Delta} \Gamma \Delta \\ \text{Ty}_{\mathbf{K} \Delta} \Gamma \underline{A} &: \equiv \text{Ty}_{\Delta} \Gamma \\ \text{Tm}_{\mathbf{K} \Delta} \Gamma A \underline{t} &: \equiv \text{Tm}_{\Delta} \Gamma A \end{aligned}$$

All structure is inherited from Δ . There is also a type substitution rule, expressing that for $\sigma : \mathbf{Sub} \Gamma \Xi$, we have $(\mathbf{K} \{\Xi\} \Delta)[\sigma] \equiv \mathbf{K} \{\Gamma\} \Delta$. This follows immediately from the above definition and the definition of type substitution, since the base inhabitants are ignored the same way on both sides of the equation. We also need to show $\mathbf{Tm} \Gamma (\mathbf{K} \Delta) \equiv \mathbf{Sub} \Gamma \Delta$. This again follows directly from the \mathbf{K} definition. From \mathbf{K} , we get

- The unit type, defined as $\mathbf{K} \bullet : \mathbf{Ty} \Gamma$.
- Direct products of Γ and Δ , defined as $\Gamma \triangleright \mathbf{K} \Delta$.
- The ability to define closed type formers as elements of \mathbf{Con} .

Universe

Similarly to what we did in Definition 32, we define \mathbf{U} as a context, and use \mathbf{K} later to get the universe as a type. $\mathbf{U} : \mathbf{Con}$ is defined to be the flcwf where objects are inner types, and morphisms are outer functions between them:

$$\begin{aligned} \text{Con}_{\mathbf{U}} &: \equiv \text{Ty}_0 \\ \text{Sub}_{\mathbf{U}} \Gamma \Delta &: \equiv \text{Tm}_0 \Gamma \rightarrow \text{Tm}_0 \Delta \\ \text{Ty}_{\mathbf{U}} \Gamma &: \equiv \text{Tm}_0 \Gamma \rightarrow \text{Ty}_0 \\ \text{Tm}_{\mathbf{U}} \Gamma A &: \equiv (\gamma : \text{Tm}_0 \Gamma) \rightarrow \text{Tm}_0 (A \gamma) \end{aligned}$$

Substitution for types and terms is defined by function composition. The empty context is defined as the inner unit type Ty_0 , and context extension $\Gamma \triangleright_{\mathbf{U}} A$ is

defined as $(\gamma : \Gamma) \times A \gamma$ using inner Σ . We can also define $\Sigma_{\mathbf{U}}$ and $\text{Id}_{\mathbf{U}}$ using inner Σ and identity.

For constant families, we don't need any additional assumption in the inner theory, since it can be defined as $\mathbf{K}_{\mathbf{U}} \{ \Gamma \} \Delta \equiv \Delta$, and $\text{Sub}_{\mathbf{U}} \Gamma \Delta \equiv \text{Tm}_{\mathbf{U}} \Gamma (\mathbf{K}_{\mathbf{U}} \Delta)$ follows immediately.

For $\mathbf{a} : \mathbf{Sub} \Gamma \mathbf{U}$, we have to define $\mathbf{El} \mathbf{a} : \mathbf{T} \Gamma$. This is given as the *displayed flowf of elements* of \mathbf{a} .

Background: from any functor $F : \mathbb{C} \rightarrow \mathbf{Set}$ we can construct the category of elements $\int F$, where objects are in $(i : |\mathbb{C}|) \times F i$ and morphisms between (i, x) and (j, y) are in $(f : \mathbb{C}(i, j)) \times (F f x \equiv y)$. If we take the second projections of components in $\int F$, we get the displayed category of elements, which lies over \mathbb{C} . We may also call this a *discrete displayed category*, in analogy to discrete categories, whose objects are elements of sets.

We extend this to flowfs in the definition of $\mathbf{El} \mathbf{a}$. With this definition, $\Gamma \triangleright \mathbf{El} \mathbf{a}$ will yield the flowf of elements of \mathbf{a} .

$$\begin{aligned} \text{Con}_{\mathbf{El} \mathbf{a}} \underline{\Gamma} &::= \text{Tm}_0 (\mathbf{a} \underline{\Gamma}) \\ \text{Sub}_{\mathbf{El} \mathbf{a}} \Gamma \Delta \underline{\sigma} &::= \mathbf{a} \underline{\sigma} \Gamma \equiv \Delta \\ \text{Ty}_{\mathbf{El} \mathbf{a}} \Gamma \underline{A} &::= \text{Tm}_0 (\mathbf{a} \underline{A} \Gamma) \\ \text{Tm}_{\mathbf{El} \mathbf{a}} \Gamma A \underline{t} &::= \mathbf{a} \underline{t} \Gamma \equiv A \end{aligned}$$

Let's check that we have all other structure as well.

- For contexts and types, the task is to exhibit elements of \mathbf{a} lying over specific base contexts and types.
- For terms and substitutions, the task is to exhibit equations which specify the action of \mathbf{a} .
- Equations between terms and substitutions are trivial because of UIP (we need to show equations between equality proofs).

We summarize below the additional structure on top of the displayed category part of $\mathbf{El} \mathbf{a}$.

- For $\bullet_{\mathbf{El} \mathbf{a}} : \text{Con}_{\mathbf{El} \mathbf{a}} \bullet$, the type can be simplified along the definition of $\text{Con}_{\mathbf{El} \mathbf{a}}$ and structure-preservation by \mathbf{a} to $\text{Tm}_0 \top_0$. Hence, $\bullet_{\mathbf{El} \mathbf{a}} \equiv \text{tt}_0$ is the unique

definition. For $\epsilon : \text{Sub}_{\mathbf{El}\mathbf{a}} \Gamma \bullet_{\mathbf{El}\mathbf{a}} \epsilon$, we have to show $\mathbf{a} \epsilon \Gamma \equiv \text{tt}_0$, which holds by the uniqueness of tt_0 .

- For $\Gamma \triangleright_{\mathbf{El}\mathbf{a}} A : \text{Con}_{\mathbf{El}\mathbf{a}} (\underline{\Gamma} \triangleright \underline{A})$, the target type unfolds to $\text{Tm}_0 (\mathbf{a} (\underline{\Gamma} \triangleright \underline{A}))$, which in turn simplifies to $\text{Tm}_0 ((\gamma : \mathbf{a} \underline{\Gamma}) \times \mathbf{a} \underline{A} \gamma)$. Since $\Gamma : \text{Tm}_0 (\mathbf{a} \underline{\Gamma})$ and $A : \text{Tm}_0 (\mathbf{a} \underline{A} \Gamma)$, we define $\Gamma \triangleright_{\mathbf{El}\mathbf{a}} A$ as (Γ, A) .
- For comprehension, we have to show the following, after simplifying types:

$$\begin{aligned} \mathbf{p} & : \mathbf{a} \underline{\mathbf{p}} (\Gamma, A) \equiv \Gamma \\ \mathbf{q} & : \mathbf{a} \underline{\mathbf{q}} (\Gamma, A) \equiv A \\ (\sigma, t) & : \mathbf{a} (\underline{\sigma}, \underline{t}) \Gamma \equiv (\Delta, A) \end{aligned}$$

For \mathbf{p} and \mathbf{q} , equations follow from preservation by \mathbf{a} . For pairing, the goal further simplifies to $(\mathbf{a} \underline{\sigma} \Gamma, \mathbf{a} \underline{t} \Gamma) \equiv (\Delta, A)$. Then, the first and second components are equal by the σ and t hypotheses.

- Assuming $A : \text{Ty}_{\mathbf{El}\mathbf{a}} \Delta \underline{A}$ and $\sigma : \text{Sub}_{\mathbf{El}\mathbf{a}} \Gamma \Delta \underline{\sigma}$, we aim to define $A[\sigma]_{\mathbf{El}\mathbf{a}} : \text{Ty}_{\mathbf{El}\mathbf{a}} \Gamma (\underline{A}[\underline{\sigma}])$. Simplifying types, $A : \text{Tm}_0 (\mathbf{a} \underline{A} \Delta)$, $\sigma : \mathbf{a} \underline{\sigma} \Gamma \equiv \Delta$ and the target type is $\text{Tm}_0 (\mathbf{a} (\underline{A}[\underline{\sigma}]) \Gamma)$, which is the same as $\text{Tm}_0 (\mathbf{a} \underline{A} (\mathbf{a} \underline{\sigma} \Gamma))$, by the preservation of $-[-]$ by \mathbf{a} . Hence, by the σ assumption, the target type is $\text{Tm}_0 (\mathbf{a} \underline{A} \Delta)$, so we give the following definition:

$$A[\sigma]_{\mathbf{El}\mathbf{a}} \equiv A$$

This is clearly functorial; moreover, substitution rules for the other type formers hold trivially.

- Term substitution is given by transitivity of equality.
- For $\text{Id}_{\mathbf{El}\mathbf{a}} t u : \text{Ty}_{\mathbf{El}\mathbf{a}} \Gamma (\text{Id } \underline{t} \underline{u})$, the goal type is $\text{Tm}_0 (\mathbf{a} (\text{Id } \underline{t} \underline{u}) \Gamma)$, hence $\text{Tm}_0 (\mathbf{a} \underline{t} \Gamma = \mathbf{a} \underline{u} \Gamma)$. This holds by $t : \mathbf{a} \underline{t} \Gamma \equiv A$ and $u : \mathbf{a} \underline{u} \Gamma \equiv A$. Reflexivity and equality reflection are trivial by UIP.
- For $A : \text{Ty}_{\mathbf{El}\mathbf{a}} \Gamma \underline{A}$ and $B : \text{Ty}_{\mathbf{El}\mathbf{a}} (\Gamma \triangleright A) \underline{B}$, we aim to define $\Sigma_{\mathbf{El}\mathbf{a}} A B : \text{Ty}_{\mathbf{El}\mathbf{a}} \Gamma (\Sigma \underline{A} \underline{B})$, hence

$$\begin{aligned} \Sigma_{\mathbf{El}\mathbf{a}} A B & : \text{Tm}_0 (\mathbf{a} (\Sigma \underline{A} \underline{B}) \Gamma) \\ \Sigma_{\mathbf{El}\mathbf{a}} A B & : \text{Tm}_0 ((A : \mathbf{a} \underline{A} \Gamma) \times \mathbf{a} \underline{B} (\Gamma, A)) \\ \Sigma_{\mathbf{El}\mathbf{a}} A B & \equiv (A, B) \end{aligned}$$

Projections and pairing proceed analogously to what we did for comprehension.

- For $K_{\mathbf{El}a} \Delta : \mathbf{Ty}_{\mathbf{El}a} \Gamma (K \underline{\Delta})$, the target type simplifies to $\mathbf{Tm}_0(a \underline{\Delta})$, hence we have $K_{\mathbf{El}a} \Delta \equiv \Delta$. For the specifying sort equation of K , we have to show

$$\mathbf{Sub}_{\mathbf{El}a} \Gamma \Delta \underline{\sigma} \equiv \mathbf{Tm}_{\mathbf{El}a} \Gamma (K_{\mathbf{El}a} \Delta) \underline{\sigma}$$

where $\underline{\sigma} : \mathbf{Sub} \Gamma \underline{\Delta}$ but at the same time $\underline{\sigma} : \mathbf{Tm} \Gamma (K \underline{\Delta})$ because of the K sort equation in the base. Fortunately, both sides simplify to $a \underline{\sigma} \Gamma \equiv \Delta$.

We still have to check $(\mathbf{El} a)[\sigma] \equiv \mathbf{El}(a \circ \sigma)$, the naturality rule for \mathbf{El} . We only have to check equality of underlying sets, \mathbf{Con} and \mathbf{Ty} formers, since terms and substitutions are equal by UIP. For underlying sets, both sides compute to the following:

$$\begin{aligned} \mathbf{Con} \Gamma & \equiv \mathbf{Tm}_0(a(\sigma \Gamma)) \\ \mathbf{Sub} \Gamma \Delta \underline{\sigma} & \equiv a(\sigma \underline{\sigma}) \Gamma \equiv \Delta \\ \mathbf{Ty} \Gamma \underline{A} & \equiv \mathbf{Tm}_0(a(\sigma \underline{A}) \Gamma) \\ \mathbf{Tm} \Gamma A \underline{t} & \equiv a(\sigma \underline{t}) \Gamma \equiv A \end{aligned}$$

Since σ also strictly preserves all structure, and we simply replace a action by the composite $a \circ \sigma$ action, it is straightforward to check that \mathbf{Con} and \mathbf{Ty} formers are also the same on both sides.

At this point, we have $\mathbf{U} : \mathbf{Con}$ and $\mathbf{El} : \mathbf{Sub} \Gamma \mathbf{U}$. Let's rename them to \mathbf{U}' and \mathbf{El}' respectively, and define the usual “open” versions:

$$\begin{aligned} \mathbf{U} : \mathbf{Ty} \Gamma & \quad \mathbf{El} : \mathbf{Tm} \Gamma \mathbf{U} \rightarrow \mathbf{Ty} \Gamma \\ \mathbf{U} & \equiv K \mathbf{U}' \quad \mathbf{El} a \equiv \mathbf{El}' a \end{aligned}$$

Identity

Assuming $t, u : \mathbf{Tm} \Gamma A$, extensional identity $\mathbf{Id} t u$ is defined as component-wise equality:

$$\begin{aligned} \mathbf{Con}_{\mathbf{Id} t u} \Gamma & \equiv t \Gamma \equiv u \Gamma \\ \mathbf{Sub}_{\mathbf{Id} t u} \Gamma \Delta \underline{\sigma} & \equiv t \underline{\sigma} \equiv u \underline{\sigma} \\ \mathbf{Ty}_{\mathbf{Id} t u} \Gamma \underline{A} & \equiv t \underline{A} \equiv u \underline{A} \\ \mathbf{Tm}_{\mathbf{Id} t u} \Gamma A \underline{t} & \equiv t \underline{t} \equiv u \underline{t} \end{aligned}$$

All other structure follows from structure-preservation of \mathbf{t} and \mathbf{u} . For the simplest example, $\bullet_{\mathbf{Id} \mathbf{t} \mathbf{u}} : \mathbf{t} \bullet \equiv \mathbf{u} \bullet$ holds because \mathbf{t} and \mathbf{u} both preserve \bullet . The rule $(\mathbf{Id} \mathbf{t} \mathbf{u})[\sigma] \equiv \mathbf{Id}(\mathbf{t}[\sigma])(\mathbf{u}[\sigma])$ is straightforward to check: we only have to look at the underlying sets, where e.g. both sides have $\mathbf{Con} \Gamma \equiv (\mathbf{t}(\sigma \Gamma) \equiv \mathbf{u}(\sigma \Gamma))$. It's also evident that $\mathbf{Tm} \Gamma (\mathbf{Id} \mathbf{t} \mathbf{u})$ is equivalent to $\mathbf{t} \equiv \mathbf{u}$, that is, we have reflexivity and equality reflection.

Inductive function type

For $\mathbf{a} : \mathbf{Tm} \Gamma \mathbf{U}$ and $\mathbf{B} : \mathbf{Tty}(\Gamma \triangleright \mathbf{El} \mathbf{a})$, we aim to define $\Pi \mathbf{a} \mathbf{B} : \mathbf{Tty} \Gamma$. This is dependent product of displayed flewfs, indexed over a *discrete* domain. Discreteness is critical: since morphisms in $\mathbf{El} \mathbf{a}$ are proof-irrelevant and invertible (because they are equations), we avoid the variance issues that preclude general Π -types in the cwf of categories [Joh02, Secion A1.5].

The direct definition would be to define underlying sets as products, indexed over corresponding components in $\mathbf{El} \mathbf{a}$:

$$\begin{aligned} \mathbf{Con}_{\Pi \mathbf{a} \mathbf{B}} \Gamma & \equiv (\gamma : \mathbf{a} \Gamma) \rightarrow \mathbf{Con}_{\mathbf{B}}(\Gamma, \gamma) \\ \mathbf{Sub}_{\Pi \mathbf{a} \mathbf{B}} \Gamma \Delta \underline{\sigma} & \equiv \{\gamma : \mathbf{a} \Gamma\} \{\delta : \mathbf{a} \Delta\} (\sigma : \mathbf{Sub}_{\mathbf{El} \mathbf{a}} \gamma \delta \underline{\sigma}) \rightarrow \mathbf{Sub}_{\mathbf{B}}(\Gamma \gamma)(\Delta \delta)(\underline{\sigma}, \sigma) \\ \mathbf{Tty}_{\Pi \mathbf{a} \mathbf{B}} \Gamma \underline{A} & \equiv \{\gamma : \mathbf{a} \Gamma\} (\alpha : \mathbf{a} \underline{A} \gamma) \rightarrow \mathbf{Tty}_{\mathbf{B}}(\Gamma \gamma)(\underline{A}, \alpha) \\ \mathbf{Tm}_{\Pi \mathbf{a} \mathbf{B}} \Gamma \mathbf{A} \underline{t} & \equiv \{\gamma : \mathbf{a} \Gamma\} \{\alpha : \mathbf{a} \underline{A} \gamma\} (t : \mathbf{Tm}_{\mathbf{El} \mathbf{a}} \gamma \delta \underline{t}) \rightarrow \mathbf{Tm}_{\mathbf{B}}(\Gamma \gamma)(\mathbf{A} \alpha)(\underline{t}, t) \end{aligned}$$

But just like in Definitions 4 and 6, we can contract the \mathbf{Sub} and \mathbf{Tm} definitions, since $\mathbf{Sub}_{\mathbf{El} \mathbf{a}} \gamma \delta \underline{\sigma} \equiv (\mathbf{a} \underline{\sigma} \gamma \equiv \delta)$ and $\mathbf{Tm}_{\mathbf{El} \mathbf{a}} \gamma \alpha \underline{t} \equiv (\mathbf{a} \underline{t} \gamma \equiv \alpha)$.

$$\begin{aligned} \mathbf{Con}_{\Pi \mathbf{a} \mathbf{B}} \Gamma & \equiv (\gamma : \mathbf{a} \Gamma) \rightarrow \mathbf{Con}_{\mathbf{B}}(\Gamma, \gamma) \\ \mathbf{Sub}_{\Pi \mathbf{a} \mathbf{B}} \Gamma \Delta \underline{\sigma} & \equiv (\gamma : \mathbf{a} \Gamma) \rightarrow \mathbf{Sub}_{\mathbf{B}}(\Gamma \gamma)(\Delta(\mathbf{a} \underline{\sigma} \gamma))(\underline{\sigma}, \text{refl}) \\ \mathbf{Tty}_{\Pi \mathbf{a} \mathbf{B}} \Gamma \underline{A} & \equiv \{\gamma : \mathbf{a} \Gamma\} (\alpha : \mathbf{a} \underline{A} \gamma) \rightarrow \mathbf{Tty}_{\mathbf{B}}(\Gamma \gamma)(\underline{A}, \alpha) \\ \mathbf{Tm}_{\Pi \mathbf{a} \mathbf{B}} \Gamma \mathbf{A} \underline{t} & \equiv (\gamma : \mathbf{a} \Gamma) \rightarrow \mathbf{Tm}_{\mathbf{B}}(\Gamma \gamma)(\mathbf{A}(\mathbf{a} \underline{t} \gamma))(\underline{t}, \text{refl}) \end{aligned}$$

With the contracted definition, \mathbf{Sub} and \mathbf{Tm} are only indexed over displayed objects, but not over displayed morphisms or terms anymore. So it is apparent that we cannot have issues with indexing variance. All structure in $\Pi \mathbf{a} \mathbf{B}$ is

pointwise inherited from \mathbf{B} . We list some examples below for definitions.

$$\begin{aligned}
\bullet_{\Pi \mathbf{a} \mathbf{B}} \gamma & \quad \equiv \bullet_{\mathbf{B}} \\
(\Gamma \triangleright_{\Pi \mathbf{a} \mathbf{B}} A) (\gamma, \alpha) & \equiv (\Gamma \gamma \triangleright_{\mathbf{B}} A \alpha) \\
\text{id}_{\Pi \mathbf{a} \mathbf{B}} \gamma & \equiv \text{id}_{\mathbf{B}} \\
(\sigma \circ_{\Pi \mathbf{a} \mathbf{B}} \delta) \gamma & \equiv \sigma \gamma \circ_{\mathbf{B}} \delta \gamma \\
A[\sigma]_{\Pi \mathbf{a} \mathbf{B}} \{\gamma\} \alpha & \equiv (A \alpha)[\sigma \gamma]_{\mathbf{B}} \\
K_{\Pi \mathbf{a} \mathbf{B}} \Delta \alpha & \equiv K_{\mathbf{B}} (\Delta \alpha)
\end{aligned}$$

For the specifying isomorphism $(\mathbf{app}, \mathbf{lam}) : \mathbf{Tm} \Gamma (\Pi \mathbf{a} \mathbf{B}) \simeq \mathbf{Tm} (\Gamma \triangleright \mathbf{El} \mathbf{a}) \mathbf{B}$, note that the difference in presentation is exactly component-wise currying and uncurrying. For instance, in $t : \mathbf{Tm} \Gamma (\Pi \mathbf{a} \mathbf{B})$, the underlying action on contexts has the following type:

$$(\underline{\Gamma} : \mathbf{Con}_{\Gamma})(\gamma : \mathbf{a} \underline{\Gamma}) \rightarrow \mathbf{Con}_{\mathbf{B}} (\underline{\Gamma}, \gamma)$$

While in $t : \mathbf{Tm} (\Gamma \triangleright \mathbf{El} \mathbf{a}) \mathbf{B}$, we have

$$((\underline{\Gamma}, \gamma) : (\underline{\Gamma} : \mathbf{Con}_{\Gamma}) \times \mathbf{a} \underline{\Gamma}) \rightarrow \mathbf{Con}_{\mathbf{B}} (\underline{\Gamma}, \gamma)$$

So \mathbf{app} and \mathbf{lam} are defined as component-wise uncurrying and currying respectively. Naturality of Π and \mathbf{app} again follows from the fact that flewf morphisms strictly preserve all structure, and substitution is component-wise function composition.

External function type

For $Ix : \mathbf{Ty}_0$ and $\mathbf{B} : \mathbf{Tm}_0 Ix \rightarrow \mathbf{Ty} \Gamma$, we define $\Pi^{\mathbf{ext}} Ix \mathbf{B} : \mathbf{Ty} \Gamma$ as the Ix -indexed direct product of a family of displayed flewfs.

$$\begin{aligned}
\mathbf{Con}_{\Pi^{\mathbf{ext}} Ix \mathbf{B}} \underline{\Gamma} & \quad \equiv (i : \mathbf{Tm}_0 Ix) \rightarrow \mathbf{Con}_{\mathbf{B} i} \underline{\Gamma} \\
\mathbf{Sub}_{\Pi^{\mathbf{ext}} Ix \mathbf{B}} \Gamma \Delta \underline{\sigma} & \equiv (i : \mathbf{Tm}_0 Ix) \rightarrow \mathbf{Sub}_{\mathbf{B} i} (\Gamma i) (\Delta i) \underline{\sigma} \\
\mathbf{Ty}_{\Pi^{\mathbf{ext}} Ix \mathbf{B}} \Gamma \underline{A} & \quad \equiv (i : \mathbf{Tm}_0 Ix) \rightarrow \mathbf{Ty}_{\mathbf{B} i} (\Gamma i) \underline{A} \\
\mathbf{Tm}_{\Pi^{\mathbf{ext}} Ix \mathbf{B}} \Gamma A \underline{t} & \equiv (i : \mathbf{Tm}_0 Ix) \rightarrow \mathbf{Tm}_{\mathbf{B} i} (\Gamma i) (A i) \underline{t}
\end{aligned}$$

All structure is defined in the evident pointwise way. $\mathbf{app}^{\mathbf{ext}}$ and $\mathbf{lam}^{\mathbf{ext}}$ are defined by component-wise flipping of function arguments.

Examples for computed semantics here maybe

4.3 Discussion of Semantics

4.3.1 Flcwfs For Free

We give a quick summary for using the semantics of FQII signatures. As input we pick a) a signature Γ b) a cwf \mathbb{C} with Σ , \top and extensional Id . Then, we interpret the signature in \mathbf{M} , thereby getting an flcwf in 2LTT. Then, we interpret that in presheaves over \mathbb{C} , and we get the flcwf whose objects are internal Γ -algebras in \mathbb{C} .

One use case is in building models of certain type theories. Usually, this starts with constructing the base cwf. But if the objects can be specified using an FQII signature, we get an flcwf for free. In some cases, we get exactly what's needed. For example, the flcwf of presheaves can be used as it is in the presheaf models of type theories.

In other cases, the flcwf that we get “for free” has to be extended in some ways. This often happens if the objects in the model have an internal notion of “equivalence” which has to be respected by types.

- In the setoid model, objects are setoids and types are displayed setoids with additional fibrancy structure [?].
- The groupoid model [HS96] is analogous; again types are displayed groupoids with fibrancy structure.
- Likewise, in the cubical set model [BCH14], types are displayed presheaves together with fibrancy structure (Kan composition).

In all these cases, the semantic objects have FQII signatures. We can interpret their flcwfs in **Set** and add fibrancy conditions. The cubical set model is presented exactly in this way in [BCH14], using displayed algebras. The groupoid model in [HS96] instead presents types as $\Gamma \rightarrow \mathbf{Gpd}$ functors, i.e. uses an indexed style instead of the displayed style.

In the indexed-style groupoid model, we get strictly functorial type substitution, just like in the displayed style. However, the displayed style appears to be a more general way to get strict substitution, as it works for every FQII theory. Again, although finitely complete categories can be always strictified to cwfs, if we ever need to perform calculations with the internal definitions of a model, the displayed style is much more compact.

4.3.2 Variations of the Semantics

In Section 4.1, we required that the inner theory has Σ , \top and extensional Id . Hence, when we interpret a signature, we again need to assume these type formers in \mathbb{C} . We used the assumed type formers in the definition of \mathbf{U} .

However, we can drop Id from the requirements on the inner theory, and likewise drop the identity type from flwfs , and the model still works. In this case we have a somewhat more general semantics. In particular, like in Section 3.5.2, we can interpret signatures in finite product categories, because \top and Σ can be derived from finite products in the constructed “simply typed” cwf . On the other hand, we get less out of the semantics. For instance, we cannot show equivalence of initiality and induction without Id .

If we want to trim down the assumptions on the inner theory to the minimum, we can make do with simply an inner cwf with no type formers at all. This implies that for each signature we can build a category of algebras, plus extra structure which does not require Σ or \top in the \mathbf{U} definition. So we may have displayed algebras, sections, and also functorial substitution for these, but we don’t have terminal algebras and total algebras (i.e. $-\triangleright_{\mathbf{U}}-$).

We could also add more type formers to the semantics. We may add *small limits* via the external function type Π^{ext} that we already have in signatures. Extending flwfs with Π^{ext} requires Π -types in the inner theory of 2LTT, hence in \mathbb{C} as well. The reason is that indexed products of algebras require functions in the underlying sorts. More concretely, in the definition of \mathbf{U} , we have to interpret

$$\Pi_{\mathbf{U}}^{\text{ext}} : (Ix : Ty_0) \rightarrow (\top m_0 Ix \rightarrow Ty_{\mathbf{U}} \Gamma) \rightarrow Ty_{\mathbf{U}} \Gamma$$

hence

$$\Pi_{\mathbf{U}}^{\text{ext}} : (Ix : Ty_0) \rightarrow (\top m_0 Ix \rightarrow \top m_0 \Gamma \rightarrow Ty_0) \rightarrow \top m_0 \Gamma \rightarrow Ty_0$$

This works if we can return an inner Π in the definition:

$$\Pi_{\mathbf{U}}^{\text{ext}} Ix B \gamma \equiv (i : Ix) \rightarrow B i \gamma$$

In this case, the “small limit” cwf semantics can be completed. We omit checking the details here. From Π^{ext} , we also recover indexed products, by using $\Pi^{\text{ext}} Ix (\lambda i. K \Gamma_i)$, where Γ_i is an indexed family of objects.

With the small limit semantics, if we want to have a simply typed interpretation, we can start with a cartesian closed \mathbb{C} .

4.3.3 Substitutions

Interpreting signatures is not the only potentially useful thing that we get out of the semantics. Each $\sigma : \text{Sub } \Gamma \Delta$ can be viewed as a free interpretation of the Δ theory in Γ , and we get a strict flwf morphism from the semantics.

Ornaments

One use case of **Sub** is to specify *ornaments* [?], i.e. ways to decorate structures with additional information, or dually, to erase parts of some structure. Ornaments differ from the usual forgetful maps, because they forget structure in *negative* position, i.e. in assumptions of construction rules.

Example 16. We assume $A : \text{Ty}_0$. We define the substitution which forgets elements of A -lists.

$$\begin{aligned} \sigma : \text{Sub} (\bullet \triangleright (Nat : \mathbf{U}) \triangleright (zero : \text{El } Nat) \triangleright (suc : Nat)) \\ (\bullet \triangleright (List : \mathbf{U}) \triangleright (nil : \text{El } List) \triangleright (cons : A \rightarrow^{\text{ext}} List \rightarrow List)) \end{aligned}$$

The map goes from numbers to lists because of the “contravariant” forgetfulness. We define σ by listing its component definitions.

$$\begin{aligned} List &: \equiv Nat \\ nil &: \equiv zero \\ cons &: \equiv \lambda^{\text{ext}} _ . \lambda n. suc\ n \end{aligned}$$

Example 17. We assume $\text{Nat}_0 : \text{Ty}_0$ with $zero_0$ and suc_0 , and define $\sigma : \text{Sub } \text{NatSig } \text{FinSig}$, where FinSig is as follows:

$$\begin{aligned} Fin &: \text{Nat}_0 \rightarrow^{\text{ext}} \mathbf{U} \\ zero &: (n : \text{Nat}_0) \rightarrow^{\text{ext}} \text{El } (Fin\ (suc_0\ n)) \\ suc &: (n : \text{Nat}_0) \rightarrow^{\text{ext}} Fin\ n \rightarrow \text{El } (Fin\ (suc_0\ n)) \end{aligned}$$

σ is defined as

$$\begin{aligned} Fin &: \equiv Nat \\ zero &: \equiv \lambda^{\text{ext}} _ . zero \\ suc &: \equiv \lambda^{\text{ext}} _ . \lambda n. suc\ n \end{aligned}$$

For a specific programming use case, if we have any recursive function defined on an “erased” type, we can convert that to a recursive function which acts on an “ornamented” type. For example, if we have some **Nat**-algebra Γ , the recursor yields a morphism from the initial algebra to Γ . We can map Γ to a list-algebra or a **Fin**-algebra, and then we can also use recursors for lists or **Fin**. Equivalently, we can map the unique morphism to Γ directly to a morphism between ornamented algebras.

Note though that a number of features and concepts from prior work on ornaments are not yet reproduced. For example, we don’t yet have an analogue of *algebraic ornaments*, which would allow us produce an ornamented signature as an *output* of a generic operation, instead of assuming it to begin with. Exploring ornaments with QII signatures could be part of future work.

Model constructions

In a broader context, ToS provides a synthetic language for specifying *model constructions*.

Example 18. For a simple example, we might want to show that constant families are equivalent to democracy in cwfs. Democracy means that for each $\Gamma : \mathbf{Con}$ there is a $\bar{\Gamma} : \mathbf{Ty} \bullet$ such that $\Gamma \simeq (\bullet \triangleright \bar{\Gamma})$ [CCD17, Section 3.1].

We can define a $\sigma : \mathbf{Sub} \mathbf{cwf}^{\mathbf{K}} \mathbf{cwf}^{\mathbf{dem}}$ which interprets democracy using constant families. It is the identity morphism on the cwf parts and interprets democracy as $\bar{\Gamma} := \mathbf{K} \Gamma$. The isomorphism $\Gamma \simeq (\bullet \triangleright \mathbf{K} \Gamma)$ follows from the specification of \mathbf{K} . We can also define a morphism $\sigma^{-1} : \mathbf{Sub} \mathbf{cwf}^{\mathbf{K}} \mathbf{cwf}^{\mathbf{dem}}$, which interprets $\mathbf{K} \Delta$ as $\bar{\Delta}[\epsilon]$. It is easy to check that σ^{-1} is indeed the inverse of σ . Thus we get an isomorphism of categories of models from the ToS semantics.

This construction is very simple, and would not be difficult to check without the ToS semantics. But it is generally not obvious that a certain mapping from models to models extends to a functor (or flcwf-morphism), so it may be helpful to work inside ToS.

Example 19. There is a simple way to show that if a type theory does not support η for Π , then function extensionality is not provable in the theory [BPT17]¹.

¹It is also possible to show unprovability of function extensionality *assuming* η for functions, but in significantly more complicated ways. To the author’s best knowledge, the set-based polynomial model is the easiest solution [?].

Assume some type theory with Σ , Π , Id and Bool , and abbreviate its signature as TT . We define a $\sigma : \text{Sub TT TT}$ which has identity action everywhere except on Π . There, we have

$$\begin{aligned}\Pi &: \equiv \lambda A B. \Pi A B \times \text{Bool} \\ \text{app} &: \equiv \lambda t. \text{app}(\text{proj}_1 t) \\ \text{lam} &: \equiv \lambda t. (\text{lam } t, \text{true})\end{aligned}$$

In short, we tag functions with a Bool value. This equips Π with “intensional” information, contradicting extensionality. If we have two functions which are pointwise equal, that only specifies that the function parts are equal, but does not say anything about the Bool tags. Hence, if we take any model of TT , we get a new model by the semantic action of σ , where function extensionality is false. Note though that the η rule also fails in the new model, so we had to drop η from the TT signature as well.

In [BPT17], this construction is presented for the special case where the starting model is initial. While it is easy to generalize to arbitrary starting models, it is less obvious to extend the construction to a functor of categories of models - which we do get for free here.

Limitations. In the finitary ToS syntax, when defining substitutions (that is, model constructions) we can only ever use assumed type constructors. For a counter-example, take the substitution in $\text{Sub MonoidSig CatSig}$ which maps a monoid to a single-object category. Assuming $\mathbf{M} : \mathbf{U}$ is the carrier set in MonoidSig , we would need to have the following:

$$\begin{aligned}\text{Obj} &: \equiv \top \\ \text{Hom} &: \equiv \lambda _ _ _ . \mathbf{M}\end{aligned}$$

But we have $\text{Obj} : \mathbf{U}$ in CatSig , so we would need to have $\top : \mathbf{U}$. In Chapter 5, we present a more expressive ToS which does include $\top : \mathbf{U}$, and it will be sufficient to rephrase a wide range of constructions as signature substitutions.

4.3.4 Recovering AMDS Interpretations

We have defined the \mathbf{M} model in a “bundled” fashion, but sometimes we will also need to refer to pieces of it. On Figure 4.1 we have a summary of the model.

	cwf					fl		
	Con	Sub	Ty	Tm	...	Σ	Id	K
cwf								
U								
Id	$-^A$	$-^M$	$-^D$	$-^S$				
Π								
Π^{ext}								

Figure 4.1: The flcwf model of the theory of signatures

On the left, the rows are labeled with components of ToS, while on the top we have components of flcwf. The individual rows can be further unfolded, as each of them contains multiple components. Likewise the Σ , **Id** and **K** columns can be unfolded. We get the whole model by filling every cell of the unfolded table with a definition. Of course, many of these cells are equations between equations, hence trivial by UIP.

This setup is very regular and convenient, because we can extract a displayed ToS model from any column, which may depend on columns to the left. The whole model is the total model of all columns. For example, the **Con** column doesn't depend on anything, so it's a plain model. The **Ty** column is displayed over **Con**. The **Tm** column depends on **Con** and **Ty**, but it does not depend on **Sub**.

From each displayed model, we get an eliminator, i.e. a family of interpretation functions. We note $-^A$, $-^M$, $-^D$ and $-^S$ in the table, but in principle we could refer to the eliminators of other columns as well. The interpretation functions can be defined in two ways:

- By separately taking the eliminators of each column, and referring to previous eliminators in each displayed model; e.g. referring to the eliminator functions $-^A$ in the definition of the **Ty** column.
- By taking the recursor for the entire model, and projecting out components from the result. E.g. we get $-^A$ by projecting out the first components of the interpretations of ToS objects.

However, the two versions coincide because of the initiality of ToS syntax.

4.4 Term Algebras

In this section we proceed with the construction of term algebras for FQII signatures, together with their recursors and eliminators. First we make two significant modifications to the setup.

First, **we drop the outer theory**, and work exclusively inside an extensional type theory. The reason is the following. The main purpose of 2LTT is to generalize the semantics of signatures. In the previous section, we presented semantics for signatures, where algebras are internal to arbitrary cwfs with Σ , \top and extensional Id . This is quite general; in particular we can interpret signatures in any finitely complete category. We also allowed dropping Id from the assumptions.

In contrast, we make a lot more assumptions in the inner theory when we develop initial term algebras; we essentially have to replicate the outer features verbatim. Thus, we gain nothing by using 2LTT, compared to working in a model of an extensional TT.

What about the term model construction for simple signatures in Section 3.5.6, why did we use 2LTT there? In that case, the inner theory was *intensional*, i.e. lacked equality reflection. So there remained an interesting distinction between the inner and outer layer, which allowed us to prove *definitional* β -rules for recursors. In contrast, here we assume inner equality reflection, so we have no distinction between propositional and definitional inner equality.

Second, **we make universe levels explicit** in the semantics and constructions. So far, we have been consistently ignoring universe levels. Now, size questions are less obvious, and quite relevant to a) ensuring the consistency of assumed induction principles b) laying groundwork for bootstrapped semantics and self-describing signatures in Section 4.5.

Universe levels are a fairly bureaucratic detail in type theories. In the following we try to be as informal as is permitted by the necessity to address and represent all the essential size problems in our subjects. In the following, we describe the new universe setup, and adapt the previously described signatures and semantics to it.

$$\begin{array}{c}
\frac{i \leq j}{\Gamma \vdash \mathbf{Set}_i \leq \mathbf{Set}_j} \qquad \frac{\Gamma, x : A \vdash B \leq B'}{\Gamma \vdash (x : A) \rightarrow B \leq (x : A) \rightarrow B'} \\
\\
\frac{\Gamma \vdash A \leq A' \quad \Gamma, x : A \vdash B \leq B'}{\Gamma \vdash (x : A) \times B \leq (x : A') \times B'} \qquad \frac{}{\Gamma \vdash A \leq A} \\
\\
\frac{\Gamma \vdash A \leq B \quad \Gamma \vdash B \leq C}{\Gamma \vdash A \leq C} \qquad \frac{\Gamma \vdash A \leq A' \quad \Gamma \vdash t : A}{\Gamma \vdash t : A'}
\end{array}$$

Figure 4.2: Rules for cumulative subtyping

4.4.1 Universes & Metatheory

We have countable Russell-style \mathbf{Set}_i universes, which are *cumulative*, meaning that any type in \mathbf{Set}_i is also an element of \mathbf{Set}_{i+1} . We use a surface syntax which is similar to Coq, where cumulativity is implicit. This contrasts the formal (“algebraic”) specification of cumulativity [Ste19, Kov21], which involves a substantial amount of explicit annotations.

Also following Coq, we have implicit *cumulative subtyping* [TS18]. In our case, this means that cumulativity commutes through basic type formers. Concretely, we have a $- \leq -$ subtyping relation on types, specified by the following rules: These rules specify subtyping for *surface syntax*; it is expected that surface syntax can be elaborated to *coercions* in a formal syntax with algebraic cumulativity.

Note that we have an *invariant* rule for function domain types. This is to match Coq and [TS18], and also because we will not need a contravariant rule in any case.

We assume that Π and Σ types return in least upper bounds of levels. For instance, assuming $A : \mathbf{Set}_i$ and $B : A \rightarrow \mathbf{Set}_j$, we have $(x : A) \rightarrow B : \mathbf{Set}_{i \sqcup j}$.

4.4.2 Signatures & Semantics

First, we parameterize the notion of ToS-model with levels.

Definition 44. For levels i and j , $\mathbf{ToS}_{i,j} : \mathbf{Set}_{i+1 \sqcup j+1}$ is the type of ToS models, defined as before, but where \mathbf{Con} , \mathbf{Sub} , \mathbf{Tm} and \mathbf{Tm} all return in \mathbf{Set}_i , and Π^{ext} abstracts over \mathbf{Set}_j .

We have that $\text{ToS}_{i,j} \leq \text{ToS}_{i+1,j}$. This follows from the rules in Figure 4.2. All underlying sets return in Set_i , which can be bumped to Set_{i+1} . The j level doesn't change, which is as expected, since Set_j appears in a negative position in the type of Π^{ext} , and has to be invariant.

Assumption. We assume that for all j , there exists $\text{syn}_j : \text{ToS}_{j+1,j}$ which supports induction. Note the level bump in the first index; this is to avoid inconsistency from type-in-type:

$$\begin{aligned} \text{Ty} &: \text{Con} \rightarrow \text{Set}_{j+1} \\ \Pi^{\text{ext}} &: (A : \text{Set}_j) \rightarrow (A \rightarrow \text{Ty } \Gamma) \rightarrow \text{Ty } \Gamma \end{aligned}$$

With Ty returning in Set_j , Π^{ext} would “contain” a Set_j , but at the same time return in a type in Set_j , and by induction we would be able to derive a Russell-like paradox [?]. Likewise, all other underlying sets must be bumped to Set_{j+1} , because of their mutual nature: contexts, terms and substitutions all “contain” types through some of their constructors.

Definition 45 (Signatures). We define $\text{Sig}_j : \text{Set}_{j+1}$ as the type of signatures where Π^{ext} may abstract over Set_j , so we have $\text{Sig}_j \equiv \text{Con}_{\text{syn}_j}$.

Definition 46 (Flwcf model). For levels i and j , we have $\mathbf{M}_{i,j} : \text{ToS}_{(i+1 \sqcup j)+1,j}$ as the model where contexts are flwfs, and objects in the flwcf are algebras. The model is defined in essentially the same way as in Section 4.2. The algebras have underlying sets in Set_i and (semantic) external functions have types in Set_j as domain. Hence, every algebra in $\mathbf{M}_{i,j}$ is in $\text{Set}_{i+1 \sqcup j}$.

Example 20. We may define NatSig as an element of Sig_0 . Then, by interpreting the signature in $\mathbf{M}_{i,0}$, we get $\text{NatSig}^A \equiv (N : \text{Set}_i) \times (N \rightarrow N) \times N$, hence $\text{NatSig}^A : \text{Set}_{i+1 \sqcup 0}$.

Notation 18. For a signature $\Gamma : \text{Sig}_j$ and level i , we may write Γ_i^A for the type of Γ -algebras with underlying sets in Set_i , which is computed by interpreting Γ in $\mathbf{M}_{i,j}$. We may use similar notation for $-^M$, $-^D$ and $-^S$.

Cumulativity of algebras. In the following, we shall assume that for $\Gamma : \text{Sig}_j$ and $i \leq i'$, we have $\Gamma_i^A \leq \Gamma_{i'}^A$. For any concrete signature Γ , this is clearly the case, but $-\leq-$ is not subject to propositional reasoning, so we cannot prove this by internal induction on signatures. We could prove that there exists a *lifting*, a $\text{Lift } \Gamma_i^A : \text{Set}_{i+1 \sqcup j}$ which is isomorphic to Γ_i^A . Instead, we take liberties, and

work as if we had actual cumulative subtyping. This seems acceptable, since by using implicit cumulativity, we are already taking the same liberty everywhere, by omitting formal lifts and isomorphisms.

4.4.3 Term Algebra Construction

We fix $\Omega : \mathbf{Sig}_j$ for some j level. We define $-^T$ by induction on \mathbf{syn}_j . In the following we write $-^A$ for $-_{j+1}^A$, i.e. the algebra interpretation where underlying sets are in \mathbf{Set}_{j+1} . Formally, we need a displayed model over \mathbf{syn}_j , but we instead present the resulting eliminator, which is perhaps easier to read. The underlying functions have the following types.

$$\begin{aligned} -^T : (\Gamma : \mathbf{Con}) \quad (\nu : \mathbf{Sub} \Omega \Gamma) &\rightarrow \Gamma^A \\ -^T : (\sigma : \mathbf{Sub} \Gamma \Delta) (\nu : \mathbf{Sub} \Omega \Gamma) &\rightarrow \Delta^T (\sigma \circ \nu) \equiv \sigma^A (\Gamma^T \nu) \\ -^T : (A : \mathbf{Ty} \Gamma) \quad (\nu : \mathbf{Sub} \Omega \Gamma) &\rightarrow \mathbf{Tm} \Omega (A[\nu]) \rightarrow A^A (\Gamma^T \nu) \\ -^T : (t : \mathbf{Tm} \Gamma A) \quad (\nu : \mathbf{Sub} \Omega \Gamma) &\rightarrow A^T \nu (A[\nu]) \equiv t^A (\Gamma^T \nu) \end{aligned}$$

We review the idea of term algebras. In any model of ToS, we might think of a $\mathbf{Sub} \bullet \Gamma$ as a Γ -algebra internal to the model. In the $-^T$ interpretation we can assume $\Omega \equiv \bullet$; this means that from any internal Γ -algebra we can extract an “external” Γ -algebra. This is possible because every sort $a : \mathbf{Tm} \Gamma \mathbf{U}$ in ToS induces an external type of terms as $\mathbf{Tm} \Gamma (\mathbf{El} a)$.

We can view the generalization from \bullet to arbitrary Ω as switching from working in the syntactic model \mathbf{syn}_j , to working in the *slice model* \mathbf{syn}_j/Ω , where contexts are Ω extended with zero or more entries. And in \mathbf{syn}_j/Ω , we have an Ω -algebra quite trivially, by taking the identity morphism $\mathbf{id} : \mathbf{Sub} \Omega \Omega^2$. Hence, term algebras arise by first taking the trivial internal algebra \mathbf{id} in \mathbf{syn}_j/Ω , then converting it to an external algebra as $\Omega^T \mathbf{id} : \Omega^A$.

Remark. We could have presented $-^T$ and slice models separately. We instead chose to merge them into the current $-^T$, since we don’t use slice models elsewhere, and we can skip their definition this way. Slice models would require the specification of *telescopes*, used to extend the base context, but this entails a fair amount of bureaucratic detail.

We explain the $-^T$ specification in the following. Term and substitution equations are given by UIP. We omit cases for substitutions and terms.

²Writing $-\mathbf{syn}_j/\Omega$ for the interpretation of syntax in the slice model, $\mathbf{Sub}_{\mathbf{syn}_j/\Omega} \bullet (\Omega^{\mathbf{syn}_j/\Omega})$ is isomorphic to, but not strictly the same as $\mathbf{Sub}_{\mathbf{syn}_j} \Omega \Omega$.

For **contexts**, we simply recurse on the entries. We use a pattern matching notation for $\mathbf{Sub} \Omega (\Gamma \triangleright A)$, since any ν with this type is uniquely determined by its first and second projections $\mathbf{p} \circ \nu$ and $\mathbf{q}[\nu]$.

$$\begin{aligned} \bullet^T \nu & \quad \quad \quad \equiv \mathbf{tt} \\ (\Gamma \triangleright A)^T(\nu, t) & \equiv (\Gamma^T \nu, A^T \nu t) \end{aligned}$$

Type substitution with $\sigma : \mathbf{Sub} \Gamma \Delta$ is as follows. This is well-typed by $\sigma^T \nu : \Delta^T (\sigma \circ \nu) \equiv \sigma^A (\Gamma^T \nu)$.

$$(A[\sigma])^T \nu t \equiv A^T (\sigma \circ \nu) t$$

For the **universe**, note that $\mathbf{U}_{j+1}^A \gamma \equiv \mathbf{Set}_{j+1}$. As we mentioned before, this is the key part when we map from internal sorts to external sets. The levels line up, since in \mathbf{syn}_j we have \mathbf{Tm} returning in \mathbf{Set}_{j+1} .

$$\begin{aligned} \mathbf{U}^T : (\nu : \mathbf{Sub} \Omega \Gamma) & \rightarrow \mathbf{Tm} \Omega \mathbf{U} \rightarrow \mathbf{T} \mathbf{y} \\ \mathbf{U}^T \nu a & \equiv \mathbf{Tm} \Omega (\mathbf{El} a) \end{aligned}$$

For \mathbf{El} , we have to define

$$(\mathbf{El} a)^T : (\nu : \mathbf{Sub} \Omega \Gamma) \rightarrow \mathbf{Tm} \Omega (\mathbf{El} (a[\nu])) \rightarrow a^A (\Gamma^T \nu)$$

but since $a^T \nu : \mathbf{Tm} \Omega (\mathbf{El} (a[\nu])) \equiv a^A (\Gamma^T \nu)$, we have

$$\begin{aligned} (\mathbf{El} a)^T : (\nu : \mathbf{Sub} \Omega \Gamma) & \rightarrow \mathbf{Tm} \Omega (\mathbf{El} (a[\nu])) \rightarrow \mathbf{Tm} \Omega (\mathbf{El} (a[\nu])) \\ (\mathbf{El} a)^T \nu t & \equiv t \end{aligned}$$

The $a^T \nu$ equation is worth noting. If we have $\nu \equiv \mathbf{id}$, the equation is $a^T \mathbf{id} : \mathbf{Tm} \Omega (\mathbf{El} a) \equiv a^A (\Omega^T \mathbf{id})$, that is, if we evaluate a signature sort in the term model $\Omega^T \mathbf{id}$, we get a type of inner terms.

For the **identity type**, we have to show that provably equal terms are evaluated to the same value in the term model.

$$(\mathbf{Id} t u)^T : (\nu : \mathbf{Sub} \Omega \Gamma) \rightarrow \mathbf{Tm} \Omega (\mathbf{Id} (t[\nu]) (u[\nu])) \rightarrow t^A (\Gamma^T \nu) \equiv u^A (\Gamma^T \nu)$$

We know by equality reflection that $t[\nu] \equiv u[\nu]$, and we also get

$$\begin{aligned} t^T \nu : A^T \nu (t[\nu]) & \equiv t^A (\Gamma^T \nu) \\ u^T \nu : A^T \nu (u[\nu]) & \equiv u^A (\Gamma^T \nu) \end{aligned}$$

from which the target equality follows. Equality reflection for inner **ld** is crucial here. It is the reason why $-^T$ works for *quotient signatures*; equality reflection is in fact the “quotient” rule which identifies provably equal terms. For a simple example, terms with type

$$\mathbf{Tm}(\bullet \triangleright (I : \mathbf{U}) \triangleright (\text{left} : \mathbf{El} I) \triangleright (\text{right} : \mathbf{El} I) \triangleright (\text{seg} : \mathbf{ld} \, l \, r)) (\mathbf{El} I)$$

are quotiented by *seg*, which is a provable equation in the context.

For the **inductive function type**, we have to convert an inner term with Π type to an outer function.

$$\begin{aligned} (\Pi a B)^T : (\nu : \mathbf{Sub} \, \Omega \, \Gamma) &\rightarrow \mathbf{Tm} \, \Omega \, (\Pi (a[\nu]) (B[\nu \circ \mathbf{p}, \mathbf{q}])) \\ &\rightarrow (\alpha : a^A (\Gamma^T \nu)) \rightarrow B^A (\Gamma^T \nu, \alpha) \\ (\Pi a B)^T \nu t &\equiv \lambda \alpha. B^T (\nu, \alpha) (t \alpha) \end{aligned}$$

This is well-typed by $a^T \nu : \mathbf{Tm} \, \Omega \, (\mathbf{El} (a[\nu])) \equiv a^A (\Gamma^T \nu)$, which allows us to consider α to be an inner term in $\lambda \alpha. B^T (\nu, \alpha) (t \alpha)$.

For the **external function type** we just recurse through the specifying isomorphism:

$$\begin{aligned} (\Pi^{\text{ext}} A B)^T : (\nu : \mathbf{Sub} \, \Omega \, \Gamma) &\rightarrow \mathbf{Tm} \, \Omega \, (\Pi^{\text{ext}} A (\lambda \alpha. (B \alpha)[\nu])) \\ &\rightarrow (\alpha : A) \rightarrow (B \alpha)^A (\Gamma^T \nu) \\ (\Pi^{\text{ext}} A B)^T \nu t &\equiv \lambda \alpha. (B \alpha)^T (\nu, \alpha) \end{aligned}$$

This concludes the definition of $-^T$.

Definition 47. For an $\Omega : \mathbf{Con}_{\text{syn}_j}$ signature, the corresponding **term algebra** is given as $\Omega^T \text{id} : \Omega_{j+1}^A$.

Remark. If we start with a signature in syn_j , then the underlying sets in the term algebra are all in \mathbf{Set}_{j+1} . Hence, the term algebra for $\mathbf{NatSig} : \mathbf{Sig}_0$ has an underlying set in \mathbf{Set}_1 . This is perhaps inconvenient, since normally we would have natural numbers in \mathbf{Set}_0 . However, we argue that this is no issue, because we are free to specify \mathbf{Set}_0 as we like. In particular, we can say that \mathbf{Set}_0 is an *empty universe*, closed under no type formers at all (or explicitly isomorphic to \perp) in which case \mathbf{Sig}_0 stands for *closed* signatures (since Π^{ext} cannot be constructed), and it is expected that any closed inductive type would be placed in \mathbf{Set}_1 . Alternatively, we could name the bottom-most universe $\mathbf{Set}_{\text{empty}}$ or \mathbf{Set}_{-1} , and start counting non-empty universes from \mathbf{Set}_0 .

4.4.4 Recursor Construction

We continue with the construction of recursors. This is not necessary, strictly speaking, since recursion is derivable from elimination, so it would suffice to only construct eliminators. We still present recursors, for the sake of matching the presentation in Chapter 2.

The goal is to construct a morphism from a term algebra to any other $\omega : \Omega^A$ algebra. However, we have to handle universe levels as well. We want to be able to eliminate from the term algebra, which was constructed at the lowest possible level, to any other universe. So far we haven't introduced a "heterogeneous" notion of morphism, between algebras at different levels. We get this from cumulativity.

- We assume $\Omega : \text{Sig}_j$, for which we already have the term algebra $\Omega^T \text{id} : \Omega_{j+1}^A$.
- We assume some $k \geq j + 1$, and an $\omega : \Omega_k^A$, the target of recursion.
- We implicitly lift $\Omega^T \text{id}$ from level $j + 1$ to level k by cumulativity, and construct a "homogeneous" morphism from the lifted term algebra to ω .

This allows us to eliminate from $\Omega^T \text{id}$ to any level. If we want to eliminate to $k \geq j + 1$, we can lift the term algebra, and use a constructed recursor. On the other hand, if we want to eliminate to $k < j + 1$, we can instead lift the target $\omega : \Omega_k^A$ algebra to $j + 1$, and again use a constructed recursor.

In general, for any $\omega : \Omega_i^A$ and $\omega' : \Omega_j^A$, the notion of heterogeneous morphism between them arises by lifting both algebras to $i \sqcup j$, and taking homogeneous morphisms between these.

Example 21. The $\text{NatSig} : \text{Sig}_0$ signature gives rise to $\text{NatSig}^T \text{id} : \text{NatSig}_1^A$. This consist of $\text{Nat} : \text{Set}_1$ together with **zero** and **suc**. Assuming a recursion principle as described above, and $\text{Bool} : \text{Set}_0$, we may define an $\text{isZero} : \text{Nat} \rightarrow \text{Bool}$ function by "downwards" elimination. We have that $(\text{Bool}, \text{true}, \lambda _ . \text{false}) : \text{NatSig}_0^A$, so by cumulativity we also have $(\text{Bool}, \text{true}, \lambda _ . \text{false}) : \text{NatSig}_1^A$, hence by recursion we get the desired morphism from $\text{NatSig}^T \text{id}$ to this model, which contains the $\text{Nat} \rightarrow \text{Bool}$ function. We can also eliminate "upwards" by lifting $\text{NatSig}^T \text{id}$ to any NatSig_i^A for $i > 1$.

We define $-^R$ by induction on syn_j . From this, we will obtain the recursor as

$\Omega^R \text{id}$.

$$\begin{aligned}
-^R : (\Gamma : \text{Con}) \quad & (\nu : \text{Sub } \Omega \Gamma) \rightarrow \Gamma^M (\nu^A (\Omega^T \text{id})) (\nu^A \omega) \\
-^R : (\sigma : \text{Sub } \Gamma \Delta) (\nu : \text{Sub } \Omega \Gamma) \rightarrow & \Delta^R (\sigma \circ \nu) \equiv \sigma^M (\Gamma^R \nu) \\
-^R : (A : \text{Ty } \Gamma) \quad & (\nu : \text{Sub } \Omega \Gamma) (t : \text{Tm } \Omega (A[\nu])) \rightarrow A^M (t^A (\Omega^T \text{id})) (t^A \omega) (\Gamma^R \nu) \\
-^R : (t : \text{Tm } \Gamma A) \quad & (\nu : \text{Sub } \Omega \Gamma) \rightarrow A^R \nu (t[\nu]) \equiv t^A (\Gamma^R \nu)
\end{aligned}$$

Maybe do an appendix with full AMDS

Let's refresh some details about the involved operations.

- For $\nu : \text{Sub } \Gamma \Delta$, we get $\nu^A : \Gamma^A \rightarrow \Delta^A$. In the semantics, ν is a functor, and ν^A is its action on objects. Analogously, for a term $t : \text{Tm } \Gamma A$, we have $t^A : (\gamma : \Gamma^A) \rightarrow A^A \gamma$, also an action on objects.
- Γ^M is the set of Γ -morphisms. $A : \text{Ty } \Gamma$ is a displayed flow in the semantics. A^M yields sets of displayed morphisms, corresponding to the semantic **Sub** component. So we have

$$A^M : A^A \gamma_0 \rightarrow A^A \gamma_1 \rightarrow \Gamma^M \gamma_0 \gamma_1 \rightarrow \text{Set}_k$$

- t^M and σ^M yield actions on morphisms. For $t : \text{Tm } \Gamma A$ and $\sigma : \text{Sub } \Gamma \Delta$, we have

$$\begin{aligned}
t^M : (\gamma^M : \Gamma^M \gamma_0 \gamma_1) & \rightarrow A^M (t^A \gamma_0) (t^A \gamma_1) \gamma^M \\
\sigma^M : (\gamma^M : \Gamma^M \gamma_0 \gamma_1) & \rightarrow \Delta^M (\sigma^A \gamma_0) (\sigma^A \gamma_1)
\end{aligned}$$

Again, we follow the “sliced” pattern that we have seen in the term model construction. Another way to view this, is that getting term algebras or recursors by direct induction on signatures is futile, since in the construction we have to refer to the *whole* Ω signature, but when we recurse inside Ω we necessarily get *smaller* signatures.

Hence, the sliced induction can be viewed as induction on arbitrary Γ signatures which are smaller than Ω , in the sense that there is a $\text{Sub } \Omega \Gamma$. Of course, $\text{Sub } \Omega \Gamma$ includes “being smaller”, but it is more general.

We look at the interpretation of type formers. Again, term and substitution equations are given by UIP, and we omit term and substitution formers. For

contexts, we again simply recurse:

$$\begin{aligned} \bullet^R \nu & \quad \quad \quad : \equiv \text{tt} \\ (\Gamma \triangleright A)^R(\nu, t) & : \equiv (\Gamma^R \nu, A^R \nu t) \end{aligned}$$

Type substitution with $\sigma : \text{Sub } \Gamma \Delta$ also follows the same pattern. The following is well-typed by $\sigma^R \nu : \Delta^R(\sigma \circ \nu) \equiv \sigma^M(\Gamma^R \nu)$.

$$(A[\sigma])^R \nu t : \equiv A^R(\sigma \circ \nu) t$$

For the **universe**, we have the following.

$$\mathbf{U}^R : (\nu : \text{Sub } \Omega \Gamma)(a : \text{Tm } \Omega \mathbf{U}) \rightarrow \mathbf{U}^M(a^A(\Omega^T \text{id}))(a^A \omega)(\Gamma^R \nu)$$

Morphisms in the semantics of \mathbf{U} are simply functions. Moreover, we have $a^T \text{id} : \text{Tm } \Omega(\text{El } a) \equiv a^A(\Omega^T \text{id})$.

$$\begin{aligned} \mathbf{U}^R : (\nu : \text{Sub } \Omega \Gamma)(a : \text{Tm } \Omega \mathbf{U}) & \rightarrow \text{Tm } \Omega(\text{El } a) \rightarrow a^A \omega \\ \mathbf{U}^R \nu a t & : \equiv t^A \omega \end{aligned}$$

Thus, we evaluate t in the ω algebra, the same way as we did in Chapter 2.

For **El**, we need to show

$$(\text{El } a)^R : (\nu : \text{Sub } \Omega \Gamma)(t : \text{Tm } \Gamma(\text{El}(a[\nu]))) \rightarrow a^M(\Gamma^R \nu)(t^A(\Omega^T \text{id})) \equiv t^A \omega$$

We have $a^R \nu : \mathbf{U}^R \nu(a[\nu]) \equiv a^M(\Gamma^R \nu)$. Hence, $\mathbf{U}^R \nu(a[\nu]) t \equiv a^M(\Gamma^R \nu) t$, and by computing \mathbf{U}^R we have $t^A \omega \equiv a^M(\Gamma^R \nu) t$. The target equation then follows by $t^T \text{id} : t^A(\Omega^T \text{id}) \equiv t$.

For **the identity type**, we need to show:

$$(\text{Id } t u)^R : (\nu : \text{Sub } \Omega \Gamma)(e : \text{Tm } \Gamma(\text{Id}(t[\nu])(u[\nu]))) \rightarrow t^M(\Gamma^R \nu) \equiv u^M(\Gamma^R \nu)$$

This follows from equality reflection on e , together with

$$\begin{aligned} t^R \nu : A^R \nu(t[\nu]) & \equiv t^M(\Gamma^R \nu) \\ u^R \nu : A^R \nu(u[\nu]) & \equiv u^M(\Gamma^R \nu) \end{aligned}$$

For the **inductive function type**, we get the following target type after unfolding $(\Pi a B)^M$:

$$\begin{aligned} (\Pi a B)^R : (\nu : \text{Sub } \Omega \Gamma)(t : \text{Tm } \Omega (\Pi (a[\nu]) (B[\nu \circ \mathbf{p}, \mathbf{q}]))) \\ \rightarrow (\alpha : a^A (\nu^A (\Omega^T \text{id}))) \rightarrow B^M (t^A (\Omega^T \text{id}) \alpha) (t^A \omega (a^M (\Gamma^R \nu) \alpha)) (\Gamma^R \nu, \text{refl}) \end{aligned}$$

We have

$$\begin{aligned} \nu^T \text{id} : \Gamma^T \nu &\equiv \nu^A (\Omega^T \text{id}) \\ a^T \nu : a^A (\Gamma^T \text{id}) &\equiv \text{Tm } \Omega (\text{El } (a[\nu])) \end{aligned}$$

Hence, $a^A (\nu^A (\Omega^T \text{id})) \equiv \text{Tm } \Omega (\text{El } (a[\nu]))$. We also have $a^R \nu : (\lambda \alpha. \alpha^A \omega) \equiv a^M (\Gamma^R \nu)$, therefore $\alpha^A \omega \equiv a^M (\Gamma^R \nu)$. With this in mind, the goal type can be rewritten as

$$\begin{aligned} (\Pi a B)^R : (\nu : \text{Sub } \Omega \Gamma)(t : \text{Tm } \Omega (\Pi (a[\nu]) (B[\nu \circ \mathbf{p}, \mathbf{q}]))) \\ \rightarrow (\alpha : \text{Tm } \Omega (\text{El } (a[\nu]))) \rightarrow B^M (t^A (\Omega^T \text{id}) \alpha) (t^A \omega (\alpha^A \omega)) (\Gamma^R \nu, \text{refl}) \end{aligned}$$

We have the following typing now:

$$B^R (\nu, \alpha) (t \alpha) : B^M ((t \alpha)^A (\Omega^T \text{id})) ((t \alpha)^A \omega) (\Gamma^T \nu, \text{refl})$$

By the action of $-^A$ on inductive function application, we have

$$B^R (\nu, \alpha) (t \alpha) : B^M (t^A (\Omega^T \text{id}) (\alpha^A (\Omega^T \text{id}))) (t^A \omega (\alpha^A \omega)) (\Gamma^T \nu, \text{refl})$$

But since $\alpha^T \text{id} : \alpha^A (\Omega^T \text{id}) \equiv \alpha$, this is exactly the target type. Therefore the definition is:

$$(\Pi a B)^R \nu t \equiv \lambda \alpha. B^R (\nu, \alpha) (t \alpha)$$

For the **external function type**, we again simply recurse through the indexing:

$$\begin{aligned} (\Pi^{\text{ext}} A B)^R : (\nu : \text{Sub } \Omega \Gamma)(t : \text{Tm } \Omega (\Pi^{\text{ext}} A (\lambda \alpha. (B \alpha)[\nu]))) \\ \rightarrow (\alpha : A) \rightarrow (B \alpha)^M (t^A (\Omega^T \text{id}) \alpha) (t^A \omega \alpha) (\Gamma^R \nu) \\ (\Pi^{\text{ext}} A B)^R \nu t \equiv \lambda \alpha. (B \alpha)^R \nu (t \alpha) \end{aligned}$$

This concludes the definition of $-^R$.

Definition 48 (Recursors). Assuming $\Omega : \text{Sig}_j$, a k level such that $k \geq j + 1$ and $\omega : \Omega_k^A$, we have $\Omega^R \text{id} : \Omega^M (\Omega^T \text{id}) \omega$ as the recursor for the term algebra.

4.4.5 Eliminator Construction

We assume $\Omega : \text{Sig}_j$ and $\omega^D : \Omega_k^D (\Omega^T \text{id})$, where $k \geq j + 1$. Again we implicitly lift the term algebra from level $j + 1$ to k . Here, ω^D is a displayed algebra over the term algebra. We seek to construct an inhabitant of $\Omega^S (\Omega^T \text{id}) \omega^D$. We define $-^E$ by induction.

Constructing eliminators is on the whole quite similar to the recursor construction. The switch from morphisms to sections is mechanical. We shall only look at **U**, **El** and **Π** here.

$$\begin{aligned} -^E : (\Gamma : \text{Con}) \quad (\nu : \text{Sub } \Omega \Gamma) &\rightarrow \Gamma^S (\nu^A (\Omega^T \text{id})) (\nu^D \omega^D) \\ -^E : (\sigma : \text{Sub } \Gamma \Delta) (\nu : \text{Sub } \Omega \Gamma) &\rightarrow \Delta^E (\sigma \circ \nu) \equiv \sigma^S (\Gamma^E \nu) \\ -^E : (A : \text{Ty } \Gamma) \quad (\nu : \text{Sub } \Omega \Gamma) (t : \text{Tm } \Omega (A[\nu])) &\rightarrow A^S (t^A (\Omega^T \text{id})) (t^D \omega^D) (\Gamma^E \nu) \\ -^E : (t : \text{Tm } \Gamma A) \quad (\nu : \text{Sub } \Omega \Gamma) &\rightarrow A^E \nu (t[\nu]) \equiv t^S (\Gamma^E \nu) \end{aligned}$$

For the **universe**, we have the following.

$$\text{U}^E : (\nu : \text{Sub } \Omega \Gamma) (a : \text{Tm } \Omega \text{U}) \rightarrow (\alpha : a^A (\Omega^T \text{id})) \rightarrow a^D \omega^D \alpha$$

By $a^T \text{id} : a^A (\Omega^T \text{id}) \equiv \text{Tm } \Omega (\text{El } a)$, we can give the following definition:

$$\begin{aligned} \text{U}^E : (\nu : \text{Sub } \Omega \Gamma) (a : \text{Tm } \Omega \text{U}) &\rightarrow (\alpha : \text{Tm } \Omega (\text{El } a)) \rightarrow a^D \omega^D \alpha \\ \text{U}^E \nu a \alpha &\equiv \alpha^D \omega^D \end{aligned}$$

In other words, we evaluate α in the ω^D displayed algebra. Let's check that this is well-typed:

$$\begin{aligned} \alpha^D &: \{\omega : \Omega^A\} (\omega^D : \Omega^D \omega) \rightarrow a^D \omega^D (\alpha^A \omega) \\ \alpha^D \omega^D &: a^D \omega^D (\alpha^A (\Omega^T \text{id})) \\ \alpha^T \text{id} &: \alpha^A (\Omega^T \text{id}) \equiv \alpha \end{aligned}$$

Thus $\alpha^D \omega^D : a^D \omega^D \alpha$. Recall that α^D can be viewed as the logical predicate interpretation of α , which expresses that α^A preserves $-^D$ predicates.

For **El**, we need to show

$$(\text{El } a)^S : (\nu : \text{Sub } \Omega \Gamma) (t : \text{Tm } \Gamma (\text{El } (a[\nu]))) \rightarrow a^S (\Gamma^E \nu) (t^A (\Omega^T \text{id})) \equiv t^D \omega^D$$

This follows from $t^T \text{id} : t^A (\Omega^T \text{id}) \equiv t$ and $a^E \nu : (\lambda t. t^D \omega^D) \equiv a^S (\Gamma^E \nu)$.

The **inductive function** interpretation is defined similarly as before:

$$\begin{aligned} (\Pi a B)^E : (\nu : \text{Sub } \Omega \Gamma) (t : \text{Tm } \Omega (\Pi (a[\nu]) (B[\nu \circ \mathbf{p}, \mathbf{q}]))) \\ \rightarrow (\alpha : \text{Tm } \Omega (\text{El } (a[\nu]))) \rightarrow B^S (t^A (\Omega^T \text{id}) \alpha) (t^D \omega^D (\alpha^D \omega^D)) (\Gamma^E \nu, \text{refl}) \\ (\Pi a B)^E \nu t \equiv \lambda \alpha. B^E (\nu, \alpha) (t \alpha) \end{aligned}$$

We make use of $\nu^T \text{id}$, $u^T \text{id}$, $a^E \nu$ and $a^T \nu$ to type-check the definition.

Interpretations for contexts and other type formers are also essentially the same as with recursors.

Definition 49 (Eliminators). Assuming $\Omega : \text{Sig}_j$, a k level such that $k \geq j + 1$ and $\omega^D : \Omega_k^D (\Omega^T \text{id})$, we have $\Omega^E \text{id} : \Omega^S (\Omega^T \text{id}) \omega^D$ as the eliminator.

Theorem 4. $\Omega^T \text{id}$ is initial when lifted to any $k \geq j + 1$ level.

Proof. $\Omega^T \text{id} : \Omega_k^A$ supports elimination by Definition 49, and elimination is equivalent to initiality by Theorem 1. \square

4.5 Levitation and Bootstrapping for Closed Signatures

When we previously introduced the ToS, we only specified the notion of model, and simply assumed that there is an evident notion of model morphism and also a notion of induction. For the theory of *closed* signatures, we can do better, because that ToS is itself a closed FQII theory. This is *levitation* [CDMM10], i.e. the situation where a ToS contains a signature for itself. Levitation is useful for bootstrapping: it shall be sufficient to specify *only the notion of model* for ToS, and notions of ToS-morphisms, initiality and induction can be computed from that. This bootstrapping process eliminates the need for either

- Assuming that the syntax of ToS already exists as a QIIT. This is not too elegant, since we are in the process of building metatheory for QII theories.
- Bootstrapping the ToS syntax as “raw” syntax, using simple inductive types, typing/conversion relations and quotients. This is very tedious and should be avoided if possible. See Section 4.6 for a discussion of this approach, although used for slightly different purposes.

In this section we describe levitation for closed signatures. The theory of closed signatures does not have Π^{ext} , but is otherwise the same as before. As we have seen, the inclusion of Π^{ext} yields a ToS which is itself infinitary, which breaks levitation. Moving to a theory of infinitary signatures will restore levitation; we revisit this in Section 5.5.

4.5.1 Models & Signatures

Since we don't have Π^{ext} , we only need a single universe level for indexing models.

Definition 50. For some i level, we have $\text{ToS}_i : \text{Set}_{i+1}$ as the type of models of ToS, where all underlying sets return in Set_i .

Definition 51 (Flwfw model). For i , we have $\mathbf{M}_i : \text{ToS}_{i+2}$ as the model where contexts are flwfw of algebras, and algebras have underlying sets in Set_i . To see how $i + 2$ checks out: if algebras contain Set_i -s, the category of algebras has a Set_{i+1} for a set of objects, and \mathbf{M}_i itself includes a category of these categories.

So far, this can be defined while only using the notion of model for ToS. What about signatures though? Previously we had that signatures are contexts in ToS *syntax*, and to talk about syntax, we need to know at least the notion of ToS model morphism.

Actually, if we only want to be able to write down signatures and interpret them in the semantics, we don't need a ToS syntax. A functional encoding suffices.

Definition 52. A **bootstrap signature** is a function which for every ToS model yields a context in that model. The type of bootstrap signatures is:

$$\text{BootSig} \equiv (i : \text{Level}) \rightarrow (M : \text{ToS}_i) \rightarrow \text{Con}_M$$

Note that this is a universe-polymorphic type. This is not an issue; universe polymorphism is a sensible feature in type theories, or alternatively we may assume that quantification over levels takes place in some outer theory.

We do not get induction on bootstrap signatures, nor do we automatically get any naturality or parametricity property.

Example 22. For NatSig , we define the expected signature, but we specify it in

an arbitrary M model instead of the syntax.

NatSig : **BootSig**

NatSig $\equiv \lambda(i : \text{Level})(M : \text{ToS}_i).$

$$(\bullet_M \triangleright_M (N : \mathbf{U}_M) \triangleright_M (\text{zero} : \mathbf{El}_M N) \triangleright_M (\text{suc} : N \Rightarrow_M \mathbf{El}_M N))$$

We might as well use the same notations for signatures as in Section 4.1, as every signature from before can be unambiguously rewritten as a bootstrap signature.

With this, we can interpret each signature in an arbitrary ToS model, by applying a signature to a model. **BootSig** _{j} can be viewed as a precursor to a Böhm-Berarducci encoding [BB85] for the theory of signatures, but we only need contexts encoded in this way, and not other ToS components. In functional programming, this style of encoding is sometimes called “finally tagless” [CKS07].

If we only want to build the 2LTT-based semantics of signatures, we are done with bootstrapping right now. In the 2LTT semantics, we never needed induction on ToS, we only needed to be able to write down signatures and interpret them in models - which we can do. Going forward, we only need to assume an inner $(\text{Ty}_0, \text{Tm}_0)$ layer with appropriate type formers, and define the flw model the same way as before.

On the other hand, if we want to consider term models, we do need a notion of induction on ToS.

Definition 53 (Signature for ToS). We define **ToSSig** : **BootSig** as the bootstrap signature for the theory of signatures. We present an excerpt from **ToSSig** below using internal notation; it should be clear that every component can be reproduced. We use **SigU** and **SigEl** to disambiguate components inside the signature from ToS components.

Con : \mathbf{U}

Sub : $\text{Con} \rightarrow \text{Con} \rightarrow \mathbf{U}$

Ty : $\text{Con} \rightarrow \mathbf{U}$

Tm : $(\Gamma : \text{Con}) \rightarrow \text{Ty } \Gamma \rightarrow \mathbf{U}$

...

SigU : $\{\Gamma : \text{Con}\} \rightarrow \mathbf{El}(\text{Ty } \Gamma)$

SigEl : $\{\Gamma : \text{Con}\} \rightarrow \text{Tm } \Gamma \text{ SigU} \rightarrow \mathbf{El}(\text{Ty } \Gamma)$

$$\begin{aligned} \Pi & : \{\Gamma : \mathbf{Con}\}(a : \mathbf{Tm} \Gamma) \rightarrow \mathbf{T}_y(\Gamma \triangleright \mathbf{SigEl} a) \rightarrow \mathbf{El}(\mathbf{T}_y \Gamma) \\ & \dots \end{aligned}$$

For each i , the interpretation of \mathbf{ToSSig} in \mathbf{M}_i yields an flewf Γ such that $\mathbf{Con}_\Gamma \equiv \mathbf{ToS}_i$, that is, objects are models of ToS at level i . This yields a model theory for ToS, which includes the notion of induction at level i .

We also know by the definition of \mathbf{ToS}_i that we have cumulativity, i.e. $\mathbf{ToS}_i \leq \mathbf{ToS}_{i+1}$ ³. Hence, we can make the following definition:

Definition 54. $M : \mathbf{ToS}_0$ supports elimination into any universe if it supports elimination when lifted by cumulativity to any \mathbf{ToS}_i .

This notion of (large) elimination is sufficient for the term algebra and eliminator constructions in Section 4.4. Thus, we were able to derive all required concepts just from the notion of model of ToS.

4.6 Reductions to Basic Type Formers

From the construction of term algebras and eliminators, we get a reduction of all QIITs to a single infinitary QIIT, namely the syntax of ToS. We spell this out:

Theorem 5. *If an extensional type theory supports syntax for $\mathbf{ToS}_{j+1,j}$, it supports initial algebras for each signature in \mathbf{Sig}_j .*

Ideally, we would like to reduce QIITs to some collection of basic type formers. The ToS syntax is far from being a basic type former, it is rather large and complicated. Therefore, the remaining job is to reduce the ToS syntax to basic type formers.

We do not attempt here to construct the entire ToS syntax as specified. The reason is the following: [LS, Section 9] shows that infinitary QIITs are not constructible from inductive types and simple quotienting with relations. We do not actually know what is the basic type former, the magic ingredient from which infinitary QIITs are actually constructible, and which is simple enough. Simplicity is key, of course; in Chapter 5 we construct infinitary QIITs from the infinitary ToS syntax, but that's again far from simple.

³For concrete bootstrap signatures we may conclude cumulativity of algebras, but we cannot conclude this universally for all bootstrap signatures, since we cannot do induction on them, and we don't even assume that they are parametric in levels.

In [AK16, Section 2.2], there is an argument that infinitary quotient inductive types extend the base type theory with constructive choice principles. We conjecture that the magic ingredient would be an infinitary QIT which expresses some kind of generic choice principle. We leave this to future work.

What we can do is to show constructions of certain fragments of the full ToS syntax. In the following, we first give a general description of QIIT constructions, then describe two specific constructions, for a) finitary inductive-inductive signatures b) closed QII signatures.

4.6.1 Finitary QIIT Constructions

The general recipe of constructing finitary QIITs from basic type formers is the following. This is more or less adapted from Streicher [Str93] and Brunerie et al. [Bru19].

1. We define the *raw* syntax, using at most inductive families, but no induction-induction. These definitions include all value constructors of the goal QIIT, but there is no indexing involved, constructors only store the raw inductive data. For example, the raw syntax of closed ToS would include the following:

$$\begin{aligned}
 \text{Con} & : \text{Set} \\
 \text{Sub} & : \text{Set} \\
 \text{Ty} & : \text{Set} \\
 \text{Tm} & : \text{Set} \\
 \bullet & : \text{Con} \\
 -\triangleright- & : \text{Con} \rightarrow \text{Ty} \rightarrow \text{Ty} \\
 \text{id} & : \text{Con} \rightarrow \text{Sub} \\
 -\circ- & : \text{Con} \rightarrow \text{Con} \rightarrow \text{Con} \rightarrow \text{Sub} \rightarrow \text{Sub} \rightarrow \text{Sub} \\
 & \dots
 \end{aligned}$$

This can be given by a simple mutual inductive definition, which can be represented as an indexed inductive family. Indexed families can be reduced in turn to indexed W-types [KvR20], which can be reduced in turn to W-types and the identity type.

2. We define typing and conversion relations on the raw syntax. For dependent type theories, the two are usually mutual: typing includes the rule which coerces terms along type conversion, and conversion is usually defined only on well-typed terms. However, it is still possible to define everything using only indexed inductive families.
3. The underlying sets are given as follows: we take raw syntactic objects which are *merely* well-formed (i.e. proofs of well-formedness are propositionally truncated, or defined in a universe of irrelevant propositions to begin with), and quotient them by conversion.
4. We show that these underlying sets support all constructors of the target QIIT: value constructors are defined using raw constructors, while equality constructors follow from conversion rules and quotienting.
5. We construct a unique morphism from the above term model to an arbitrary model of the QII theory. This usually requires several steps. One approach is to first define by induction on raw syntax a family of *partial functions* into the assumed model, then separately show that these functions are total on well-typed input. The separation is necessary because the induction principle for the raw syntax is *too weak*: it cannot express the inductive-inductive indexing dependencies which would be required to construct the morphism in one go. For instance, if we have the QIIT syntax for ToS, and we have some displayed model A over the syntax, the eliminator contains the following:

$$\begin{aligned} \text{Con}^S &: (\Gamma : \text{Con}) \rightarrow \text{Con}_A \Gamma \\ \text{Sub}^S &: (\Gamma \Delta : \text{Con})(\sigma : \text{Sub } \Gamma \Delta) \rightarrow \text{Sub}_A (\text{Con}^S \Gamma) (\text{Con}^S \Delta) \sigma \end{aligned}$$

But with the raw syntax, we can only eliminate using a displayed model of the raw syntax, and the eliminator contains the following:

$$\begin{aligned} \text{Con}^S &: (\Gamma : \text{Con}) \rightarrow \text{Con}_A \Gamma \\ \text{Sub}^S &: (\sigma : \text{Sub}) \rightarrow \text{Sub}_A \sigma \end{aligned}$$

Lastly, we show that the constructed morphism is unique. This is done by induction on raw syntax, and is generally possible in just one elimination.

Note that the above recipe permits a large number of design variations. Some examples:

- We may omit fields from raw syntax which are fully determined by type indices. This may make subsequent work easier or harder depending on particulars.
- We may start from a *well-scoped* raw syntax, if there is a notion of scoping in the goal QIIT. In general, we may start from some kind of partially raw syntax, which is well-typed to some extent. This extent is bounded by what is expressible only using indexed inductive families but not induction-induction.
- We may move along a spectrum of “paranoia” in the specification of well-typing [Win20, Section 9.2]. A paranoid typing rule assumes the well-formedness of everything involved, for example assumes the well-formedness of a context Γ before it assumes well-formedness of a type in Γ . In contrast, an “economic” specification tries to make the minimum necessary assumptions, relying on admissibility properties. It’s possible that well-formedness of Γ is derivable from the well-formedness of a type in Γ , so the assumption can be dropped.

However, if we omit too much, then some other admissibility properties may break! Design decisions along the paranoia spectrum are often all tangled up like this; hence the name “paranoid”, which probably stems from the anxiety of breaking things by making too much shortcuts.

- Instead of using partial maps from raw syntax to the the assumed model in step 5, we may define *relations* between *well-formed* raw syntax and the given model, and later show that these relations are functional. This seems to be a technically easier approach. The reason is that we do not have decidable definedness of the partial maps, which makes them more complicated. A decidable defined partial function has type $A \rightarrow \mathbf{Maybe} B$. For any $a : A$ we can look at whether the function is defined on it. A more general partial function has type $A \rightarrow ((P : \mathbf{Prop}) \times (P \rightarrow B))$. If we forget about the **Prop**-ness of P for the time being, we can equivalently have a relation $A \rightarrow B \rightarrow \mathbf{Set}$ instead. This is a more “indexed” definition compared to the “fibred” presentation with $P : \mathbf{Prop}$, and indexed presentations in type theory usually enjoy more definitional computation rules - this is also the reason why displayed algebras are better-behaved computationally than fibred algebras.

It should be apparent that constructing QIITs is tedious, and especially so for large QIITs like type theories. Hence, it is best if we do it just once, for a theory of signatures from which every other QIIT can be constructed.

Connection to the initiality conjecture

The initiality conjecture was made by Voevodsky [Voe17], and it is essentially the conjecture that the above construction (“initiality construction”) can be carried out in sufficient formal detail for “usual” type theories.

There has been much debate about the merits of initiality constructions. See [Con19] for a hub of such discussions. On one hand, some people believed that the initiality construction is essential for reconciling the usage of raw syntax and categorical notions of models. On the other hand, some people dismissed the initiality construction as a pointless exercise, considering the categorical syntax to be the actual syntax, and raw syntax as merely notation for that. The author of this thesis is of a somewhat different opinion than either of the above.

First, as a moral justification for the usage of raw syntax, the initiality construction is indeed mostly pointless. That’s because *elaboration* comprises the true justification for that. Elaboration is the effective algorithm which converts raw syntax to “core syntax”, i.e. typed categorical syntax. Given a piece of raw syntax, even if I have done the initiality construction, I have no effective way of learning which core syntactic object it corresponds to! The elaboration literature is *all about* practical justifications for using certain raw syntaxes, and it comes with established ways to talk about power and correctness of elaboration algorithms.

Second, there is a different motivation for the initiality construction: *foundational minimalism*, the reduction of a complicated QIIT to basic type formers. Elaboration merely assumes that a categorical core syntax already exists, as the target of elaboration, but it is orthogonal to the construction of the core syntax. If we have elaboration, we may still want to show a reduction of the core syntax, but now we are free to perform this construction in whatever way is the easiest. We don’t have to construct the QIIT out of a raw syntax which is intentionally close to the raw syntax that we use in practice! In the author’s opinion, a great deal of confusion arises from the conflation of the two different motivations for the initiality construction.

As to which way of construction is easiest: the author does not know of any

truly easy way, but this thesis shows that we only have to do it once, for a theory of signatures, and then we can construct all other QIITs from that in a generic way. In particular, almost all type theories in the wild are finitary closed QII theories (with the notable exceptions of our ToS-es), so if we can construct closed signatures, we can construct initial models of almost all type theories.

What about generic ways to formalize elaboration algorithms? This seems to be a lot more difficult. To the author's knowledge there have not been notable research in this area. Decidability of conversion is already very hard to analyze in a generic way, and the simplest possible bidirectional elaboration algorithms rely on decidable conversion. To formalize practically realistic elaboration (i.e. elaboration which includes unification) is yet more difficult.

4.6.2 Reduction of Finitary Inductive-Inductive Types

This section is based on the author's joint work with Kaposi and Lafont [KKL19]. The core idea is the following: a certain fragment of ToS can be constructed in a far simpler way than what we described in Section 4.6.1, with fewer assumptions in the ambient theory. We call this fragment the theory of *finitary inductive-inductive* signatures. This theory has the following type formers (on the top of the base cwf):

- The \mathbf{U} universe with \mathbf{El} .
- Inductive function type Π , but without \mathbf{lam} , and thus without $\beta\eta$ -rules.
- External function type Π^{ext} , but again without $\mathbf{lam}^{\text{ext}}$.

Concretely, this ToS is tuned so that

1. No quotients are required in its construction.
2. The generic term model construction still goes through for every signature in the ToS.

We explain in the following. First, the equational theory of ToS only specifies substitution, but it contains no computation rules for type formers. Thus, ToS is a theory of *neutral terms* and substitutions. This allows us to define a raw syntax which only includes normal forms, and to define substitution as recursive functions acting on normal forms. This trivializes the conversion relation: conversion is simply propositional equality of raw terms. Thus, there is no need to

quotient by conversion. Note that our raw syntax is infinitary, because we have to represent the branching in Π^{inf} . This is fine though: we only run into the issue of the missing “choice” principle if we try to mix quotients and infinite branching. Without quotients, infinite branching is not an issue.

Second, we do not include an identity type in ToS. This blocks the other way for quotients to enter the picture. With identity types, the generic term model construction relies on equality reflection in ToS. But when we construct ToS syntax, the only way to show equality reflection is to quotient raw syntax by internally provable equalities.

Third, it remains to check that the generic term model construction works with the pared-down ToS. We only need to check that the omission of **lam** and **lam**^{ext} does not mess things up. Looking at Sections 4.4.3 and 4.4.5, we see that they don’t: the interpretations of Π and Π^{ext} only require applications in ToS, not abstractions.

Remark. Although we have not yet talked about infinitary signatures, we can give a short summary why the current construction fails to work for their ToS. The generic term algebra construction in Section 5.4.1 for infinitary signatures relies on there being both **lam** and **app** for “infinitary” function types, with $\beta\eta$ -rules. This makes the equational theory of ToS non-trivial, so quotients are necessary in the construction of the syntax. However, this requires mixing quotients and infinite branching, and in this case we do not even know how to construct the syntax.

We summarize the construction of the ToS syntax below. We refer the reader to [KKL19] for details.

1. We define raw syntax by mutual induction. Substitution are in normal form: they are just lists of raw terms. Variables are also normalized as de Bruijn indices. We define the action of substitution by recursion on raw syntax. In [KKL19], raw syntax is not well-scoped, and substitution is partial, but it would be also possible to start from well-scoped raw syntax.
2. We inductively define well-formedness relations for contexts, substitutions, types and terms, and show by induction on raw syntax that well-formedness is propositional, i.e. proof-irrelevant. Alternatively, we could have defined well-formedness by recursion on raw syntax.
3. We construct a term model of ToS from well-formed raw syntax. All equa-

tions in the model are provable from the properties of recursive substitution on raw terms.

4. We pick a ToS model, and inductively define a family of relations between the term model and the given model, which define the function graphs of the model morphism that we aim to define. Then we show in order:
 - (a) Right-uniqueness of the relation, by induction on well-formedness derivations.
 - (b) Stability of the relation under substitution.
 - (c) Left-totality of the relation, by induction on well-formedness derivations.

We then define the actual model morphism using the functionality of the relation.

5. For the uniqueness of the constructed morphism, we exploit right-uniqueness of the relation: it is enough to show that any other model morphism maps syntactic input to related semantic output.

This construction is formalized in Agda; see [KKL19]. It uses indexed inductive families, UIP, function extensionality, and equality reflection in the form of Agda rewrite rules, although the latter could be in principle omitted from the formalization. Thus, it follows that any model of ETT with inductive families supports finitary inductive-inductive types.

4.6.3 Reduction of Closed QIITs

For closed QIITs, there is unfortunately no direct formalization which constructs the ToS. There is one though which is *close enough*, by Menno de Boer and Guillaume Brunerie [BdB20], see also De Boer’s thesis [dB20]. This constructs a type theory with the following features:

- A contextual category for base (instead of a cwf).
- Countable predicative universes.
- \mathbb{N} , Σ , Π , \top , \perp , $-+-$ and intensional Id .

The construction follows the 1-5 steps that we described previously in Section 4.6.1. It makes the following assumptions:

- A universe of strict propositions **Prop**. Every type in this universe enjoys definitional proof-irrelevance. This **Prop** is used to define partial functions and well-formedness relations.
- Function extensionality.
- Propositional extensionality for **Prop**.
- Quotients by relations valued in **Prop**.
- Indexed inductive families returning in **Set** or in **Prop**.

Notably, UIP is not assumed. It appears that the irrelevance of equalities in **Prop** is sufficient to obviate UIP.

It is very plausible that this construction can be adapted to our theory of closed QII signatures. De Boer and Brunerie construct a complicated open finitary QIIT, while ours is a fairly similar closed QIIT, with fewer and more restricted type formers. The openness comes from the use of contextual categories, which involve indexing by external natural numbers. Contextuality does not make much difference in the construction though, since raw syntax is always contextual by the inductive nature of raw contexts.

Hence, it is safe to say that any model of a type theory which supports the assumptions of De Boer and Brunerie, also supports all closed QIITs.

CHAPTER 5

Infinitary Quotient Inductive-Inductive Signatures

In this chapter we present another theory of signatures, for *infinitary quotient inductive-inductive* signatures. As we will see, the reason for considering the finitary and infinitary cases separately is that they support different semantics.

First, we specify signatures and define semantics in 2LTT. Then, like in the previous chapter, we switch to a extensional TT setting and look at term algebras and related constructions.

5.1 Theory of Signatures

Metatheory. We work in 2LTT. We assume the following type formers in the inner theory: \top , Σ , extensional identity $- = -$ and Π . Note that Π is an extra assumption compared to what we had in the finitary case.

Definition 55. A **model of the theory of signatures** consists of the following.

- A **cwf** with underlying sets Con , Sub , Ty and Tm , all returning in the outer Set universe of 2LTT.
- A **Tarski-style universe** \mathcal{U} with decoding El . \mathcal{U} is closed under the following type formers:
 - The **unit type** \top .
 - **Σ -types** $\Sigma : (a : \text{Tm } \Gamma \mathcal{U}) \rightarrow \text{Tm } (\Gamma \triangleright \text{El } a) \mathcal{U} \rightarrow \text{Tm } \Gamma \mathcal{U}$, with specifying isomorphism

$$(\text{proj}, -, -) : \text{Tm } \Gamma (\text{El } (\Sigma a b)) \simeq (t : \text{Tm } \Gamma (\text{El } a)) \times \text{Tm } \Gamma (\text{El } (b[\text{id}, t]))$$

- **Extensional identity** $\text{Id} : \text{Tm } \Gamma (\text{El } a) \rightarrow \text{Tm } \Gamma (\text{El } a) \rightarrow \text{Tm } \Gamma \text{U}$, specified by $(\text{reflect}, \text{refl}) : \text{Tm } \Gamma (\text{El } (\text{Id } t u)) \simeq (t \equiv u)$.
- **Infinitary functions** $\Pi^{\text{inf}} : (A : \text{Ty}_0) \rightarrow (A \rightarrow \text{Tm } \Gamma \text{U}) \rightarrow \text{Tm } \Gamma \text{U}$, specified by $(\text{app}^{\text{inf}}, \text{lam}^{\text{inf}}) : \text{Tm } \Gamma (\Pi^{\text{inf}} A b) \simeq ((\alpha : A) \rightarrow \text{Tm } \Gamma (\text{El } (b \alpha)))$.
- An **inductive function type** $\Pi : (a : \text{Tm } \Gamma \text{U}) \rightarrow \text{Ty } (\Gamma \triangleright \text{El } a) \rightarrow \text{Ty } \Gamma$, specified by $(\text{app}, \text{lam}) : \text{Tm } \Gamma (\Pi a B) \simeq \text{Tm } (\Gamma \triangleright \text{El } a) B$.
- An **external function type** $\Pi^{\text{ext}} : (A : \text{Ty}_0) \rightarrow (A \rightarrow \text{Ty } \Gamma) \rightarrow \text{Ty } \Gamma$, specified by $(\text{app}^{\text{ext}}, \text{lam}^{\text{ext}}) : \text{Tm } \Gamma (\Pi^{\text{ext}} A B) \simeq ((\alpha : A) \rightarrow \text{Tm } \Gamma (B \alpha))$.

Once again we assume that an initial model for ToS exists, and a signature is a context in the initial model.

Notation 19. We employ the same notations for signatures as in Section 4.1. In addition to that, we have the usual internal notation for \top and Σ , and we write $(x : A) \rightarrow^{\text{inf}} B$ for Π^{inf} and λ^{inf} for lam^{inf} .

Let's do a comparison to the finitary case. First, the new signatures do not support sort equations, since there is no identity type for arbitrary terms, only for terms with types in U . Second, the universe is not empty anymore, it supports \top , Σ and the infinitary function type Π^{inf} , which can be viewed as an analogue of Π^{ext} inside U . We look at example signatures.

Example 23. Infinitary constructors can be given with Π^{inf} . A classic example is W-types. Assuming $S : \text{Ty}_0$ and $P : S \rightarrow \text{Ty}_0$, we have the following signature for P -branching well-founded trees:

$$\begin{aligned} W & : \text{U} \\ \text{sup} & : (s : S) \rightarrow^{\text{ext}} (P s \rightarrow^{\text{inf}} W) \rightarrow \text{El } W \end{aligned}$$

Note that since $P s \rightarrow^{\text{inf}} W$ is in U , it can appear on the left side of \rightarrow . Hence the naming “infinitary”: if $P s$ is an infinite type, **sup** branches with an infinite number of inductive subtrees. Of course, finitary branching can be also expressed with Π^{inf} , but that use case was already possible with finitary signatures, by iterating \rightarrow finite times.

Example 24. Equations can appear as assumptions now. The simplest example

would be set truncation for some $A : \mathbf{Ty}_0$:

$$\begin{aligned} |A|_0 & : \mathbf{U} \\ \text{embed} & : A \rightarrow^{\text{ext}} \mathbf{El} |A|_0 \\ \text{trunc} & : (x\ y : |A|_0)(p\ q : \mathbf{ld}\ x\ y) \rightarrow \mathbf{El} (\mathbf{ld}\ p\ q) \end{aligned}$$

However, this ends up being redundant in our semantics, since we assume UIP, and every semantic underlying type will be a set. Does this mean that recursive equations are useless? We don't think so. In the specification of cubical type theories, there are *boundary conditions* which can be given as \mathbf{ld} assumptions [CCHM17,AHW16,AHH18]. Also, it seems that these conditions cannot be easily contracted away, or expressed equivalently without recursive equation. For an example of contraction, the signature

$$\bullet \triangleright (A : \mathbf{U}) \triangleright (c_1 : \mathbf{El}\ A) \triangleright (c_2 : (x : A) \rightarrow \mathbf{ld}\ x\ c_1 \rightarrow \mathbf{El}\ A)$$

can be rewritten to the equivalent

$$\bullet \triangleright (A : \mathbf{U}) \triangleright (c_1 : \mathbf{El}\ A) \triangleright (c_2 : \mathbf{El}\ A)$$

signature. However, we can't mechanically eliminate the \mathbf{ld} from the following signature.

$$\begin{aligned} A & : \mathbf{U} \\ B & : A \rightarrow \mathbf{U} \\ b_1 & : A \rightarrow \mathbf{El}\ B \\ b_2 & : A \rightarrow \mathbf{El}\ B \\ & \dots \\ a & : (x\ y : A) \rightarrow \mathbf{ld}\ (b_1\ x)\ (b_2\ y) \rightarrow \mathbf{El}\ A \end{aligned}$$

Whether we can reformulate a without the \mathbf{ld} condition depends on what kind of equational theory we specify for B in the omitted parts of the signature.

Example 25. All theories of signatures that we discussed so far, have (infinitary) signatures.

For finitary signatures, the ToS is itself infinitary because of Π^{ext} . We assume

an universe U_0 in Ty_0 . In the signature, we have

$$\begin{aligned} Con &: U \\ Ty &: Con \rightarrow U \\ \Pi^{\text{ext}} &: \{\Gamma : Con\} \rightarrow (A : U_0) \rightarrow^{\text{ext}} (A \rightarrow^{\text{inf}} Ty \Gamma) \rightarrow \text{El}(Ty \Gamma) \end{aligned}$$

In the signature for infinitary ToS, we have

$$\begin{aligned} Univ &: \{\Gamma : Con\} \rightarrow Ty \Gamma \\ \Pi^{\text{inf}} &: \{\Gamma : Con\} \rightarrow (A : U_0) \rightarrow^{\text{ext}} (A \rightarrow^{\text{inf}} Tm \Gamma Univ) \rightarrow Tm \Gamma Univ \end{aligned}$$

Remark. When we will take the semantics of the above signature, we won't exactly get back the theory of signatures that we're using right now. We have ToS in 2LTT now, but the semantics is in the inner theory. What we can do though, is to assume that the inner theory is also a 2LTT. Then we might assume that the inner theory of *that* is again a 2LTT, and so on. This is a possible (and quite natural) generalization of 2LTT to n-level type theory. In this setting, one round of self-description requires a bumping of levels in the sense of n-level TT. In this thesis we don't explore this, instead we use a more conventional universe hierarchy in an extensional TT, to investigate self-description.

Example 26. We have seen in Example 15 that Ty_0 -valued presheaves have finitary signatures. With infinitary signatures, we can also cover monads on Ty_0 . We assume a universe $U_0 : Ty_0$.

$$\begin{aligned} M &: U_0 \rightarrow U \\ map &: (A \rightarrow B) \rightarrow^{\text{ext}} M A \rightarrow \text{El}(M B) \\ map_{id} &: \text{El}(\text{Id}(map \text{ id } x) x) \\ map_{\circ} &: \text{El}(\text{Id}(map(f \circ g) x)(map f (map g x))) \\ return &: A \rightarrow^{\text{ext}} \text{El}(M A) \\ bind &: M A \rightarrow (A \rightarrow^{\text{inf}} M B) \rightarrow \text{El}(M B) \end{aligned}$$

We rely on \rightarrow^{inf} to specify binding. The *join*-based specification would not work, since $M(M A)$ is not valid in signatures. The above signature can be helpful for deriving some of the metatheory of Dijkstra monads [MAA⁺19, Section 5].

In the 2LTT-based semantics, we will get $M : U_0 \rightarrow Ty_0$, which is not quite an endofunctor, but it would be straightforward to refine the semantics so that

we can pick universes for underlying types of algebras, and have $M : \mathbf{U}_0 \rightarrow \mathbf{U}_0$. Alternatively, we will also have the ETT-based semantics, where universe levels are explicitly handled.

Example 27. It is worth to note that every set-truncated higher inductive type from the Homotopy Type Theory book [Uni13] is covered. This includes

- The cumulative hierarchy of sets [Uni13, Section 10.5].
- Cauchy real numbers [Uni13, Section 11.3].
- Surreal numbers [Uni13, Section 11.6].

5.2 Semantics

5.2.1 Overview

As we mentioned, we need a different semantics for infinitary signatures. First, we look at why the previous semantics fails. We try to model signatures again as flowfs, and morphisms as strict flowf-morphisms. The simplest point of failure is the interpretation of the unit type $\top : \mathbf{Tm} \Gamma \mathbf{U}$.

In the semantics, this is the same as defining $\top : \mathbf{Sub} \Gamma \mathbf{T}_{\mathbf{y}_0}$, where $\mathbf{T}_{\mathbf{y}_0}$ is the flowf of inner types. The only sensible definition here is the functor which is constantly \top_0 . But this does not strictly preserve comprehension or the finite limit type formers. If we have

$$\begin{aligned} \top &: \mathbf{Con}_{\Gamma} \rightarrow \mathbf{T}_{\mathbf{y}_0} \\ \top \Gamma &:= \top_0 \end{aligned}$$

then we have $\top (\Gamma \triangleright_{\Gamma} A) \equiv \top_0$, but $\top \Gamma \triangleright_{\mathbf{T}_{\mathbf{y}_0}} \top A \equiv \top_0 \times \top_0$. Thus, $\top_0 \not\equiv \top_0 \times \top_0$, but of course $\top_0 \simeq \top_0 \times \top_0$.

Let's look at $\mathbf{\Pi}^{\text{inf}} : (A : \mathbf{T}_{\mathbf{y}_0}) \rightarrow (A \rightarrow \mathbf{Tm} \Gamma \mathbf{U}) \rightarrow \mathbf{Tm} \Gamma \mathbf{U}$ as well, since that's a more interesting new feature than the unit type. The only viable definition is to take the A -indexed product of $\mathbf{Sub} \Gamma \mathbf{T}_{\mathbf{y}_0}$ morphisms, so we map objects of Γ to function types:

$$\mathbf{Con}_{\mathbf{\Pi}^{\text{inf}} \mathbf{A} \mathbf{b}} \Gamma := (\alpha : A) \rightarrow \mathbf{Con}_{\mathbf{b} \alpha} \Gamma$$

But now we have

$$(\Pi^{\text{inf}} \mathbf{A} \mathbf{b}) \bullet_{\Gamma} \equiv (\alpha : A) \rightarrow \text{Con}_{b\alpha} \bullet_{\Gamma} \equiv A \rightarrow \top_0$$

Hence, $(\Pi^{\text{inf}} \mathbf{A} \mathbf{b}) \bullet_{\Gamma} \not\equiv \top_0$, although $(\Pi^{\text{inf}} \mathbf{A} \mathbf{b}) \bullet_{\Gamma} \simeq \top_0$.

Intuitively, if \mathbf{U} has no type formers, the terms in \mathbf{U} are neutral, i.e. variables applied to zero or more neutral terms. But variables in the semantics simply project out components from iterated Σ -types. For example, the action of $\mathbf{q} : \mathbf{Tm}(\Gamma \triangleright \mathbf{A})(\mathbf{A}[\mathbf{p}])$ on objects, types, morphisms and terms is given by taking second projections. Since all structure in $\Gamma \triangleright \mathbf{A}$ is given by pairing things, \mathbf{q} strictly preserves all structure, and the same goes for all variables.

Substitutions and terms in the finitary ToS are only allowed to freely reshuffle structure. We can forget, duplicate, or permute signature entries, or build neutral expressions from assumptions. In contrast, the infinitary ToS allows us to take small limits of assumptions, using \top , Σ , \mathbf{Id} and Π^{inf} to build new inhabitants of \mathbf{U} . We summarize the process of getting the new semantics:

1. Strict structure-preservation for type formers in \mathbf{U} generally fails, but they still preserve structure up to isomorphism.
2. Hence, we switch from strict flowf-morphisms to weak ones, which preserve \bullet , comprehension and fl-structure weakly.
3. However, in the finitary case we often relied on transporting along preservation equations. We need to recover transports along isomorphism.
4. Hence, we extend semantic types from displayed flowfs to *isofibrations*, which support the required transports.
5. However, this rules out sort equations, because they are not stable under isomorphisms. For example, for sets A, B, C such that $A \simeq B$ and $A \simeq C$, it's not necessarily the case that $B \equiv C$.

Univalent semantics

The isofibrant semantics will turn out to be significantly more technical than the strict semantics. Instead of working with isofibrations in an extensional setting, could we work with univalent structures in homotopy type theory? In other words, work with univalent categories of algebras, and univalent displayed categories

over them [AL19]. A major benefit of the univalent setting is that we would get a *structure identity principle* [?] out of the semantics, which says that for algebras, isomorphism is the same as equality.

However, it appears that univalent *cwfs* are overall yet more technical to handle than isofibrations. In an univalent *cwf*, objects and types are generally *h*-groupoids, so we would have groupoids of algebras instead of sets of algebras. This implies that *type equalities* are between groupoids, so they need to be coherent, if we want them to be well-behaved. Hence, \mathbf{Ty} is not an 1-presheaf over contexts, but rather a $(2, 1)$ -presheaf.

Alternatively, we could simplify the task by only constructing univalent categories of algebras, and skipping the family structure (and *fl*-structure). This would be the minimum amount of effort that would yield the structure identity principle.

Both of these would be interesting to check in future work. As a third alternative, instead of stopping at set-truncated algebras in HoTT, we might as well consider types at arbitrary *h*-level, and construct $(\omega, 1)$ -categories of algebras. This comprises a semantics of higher inductive-inductive signatures. We do not present a full higher-categorical semantics in this thesis; we only present a fragment of it in Chapter 6.

Refer to Sattler & Capriotti in HIIT chapter, or here?

5.2.2 Model of the Theory of Signatures

In the following we present a model of ToS. We call it \mathbf{M} , and like before, we use **bold** font to refer to components of \mathbf{M} .

Contexts

$\Gamma : \mathbf{Con}$ is again an *flcwf*, but with a minor change: \mathbf{K} is not strict anymore, so we have $(\mathbf{app}_{\mathbf{K}}, \mathbf{lam}_{\mathbf{K}}) : \mathbf{Tm} \Gamma (\mathbf{K} \Delta) \simeq \mathbf{Sub} \Gamma \Delta$. As we will see shortly, $\mathbf{A}[\sigma]$ does not support strict displayed \mathbf{K} anymore, hence the change.

Substitutions

$\sigma : \mathbf{Sub} \Gamma \Delta$ is a *weak flcwf-morphism*, that is, a functor between underlying categories, which also maps types to types and terms to terms, and satisfies the following mere properties:

1. $\sigma(A[\sigma]) \equiv (\sigma A)[\sigma \sigma]$
2. $\sigma(t[\sigma]) \equiv (\sigma t)[\sigma \sigma]$
3. The unique map $\epsilon : \text{Sub}(\sigma \bullet) \bullet$ has a retraction.
4. Each $(\sigma p, \sigma q) : \text{Sub}(\sigma(\Gamma \triangleright A))(\sigma \Gamma \triangleright \sigma A)$ has an inverse.

In short, σ preserves substitution strictly and preserves empty context and context extension up to isomorphism. We notate the evident isomorphisms as $\sigma_\bullet : \sigma \bullet \simeq \bullet$ and $\sigma_\triangleright : \sigma(\Gamma \triangleright A) \simeq \sigma \Gamma \triangleright \sigma A$. Our notion of weak morphism is the same as in [BCM⁺20], when restricted to cwfs.

Theorem 6. *Every $\sigma : \text{Sub } \Gamma \Delta$ preserves fl-structure up to type isomorphism. That is, we have*

$$\begin{aligned}\sigma_\Sigma &: \sigma(\Sigma A B) \simeq \Sigma(\sigma A)((\sigma B)[\sigma_\triangleright^{-1}]) \\ \sigma_K &: \sigma(K \Delta) \simeq K(\sigma \Delta) \\ \sigma_{\text{Id}} &: \sigma(\text{Id } t u) \simeq \text{Id}(\sigma t)(\sigma u)\end{aligned}$$

These are all natural in the following sense: for $\sigma : \text{Sub}_\Gamma \Gamma \Delta$, if we have σ_Σ as a type isomorphism in $\sigma \Delta$, if we reindex it by σ , we get σ_Σ as a type isomorphism in $\sigma \Gamma$. The same holds for σ_K and σ_{Id} .

Moreover, σ preserves all term and substitution formers in the fl-structure. For example, $\sigma(\text{proj1 } t) \equiv \text{proj1}(\sigma_\Sigma[\text{id}, \sigma t])$.

Proof. For σ_Σ , we construct the following context isomorphism:

$$\begin{aligned}(\sigma \Gamma \triangleright \sigma(\Sigma A B)) &\simeq (\sigma \Gamma \triangleright \sigma A \triangleright (\sigma B)[\sigma_\triangleright^{-1}]) \\ &\simeq (\sigma \Gamma \triangleright \Sigma(\sigma A)((\sigma B)[\sigma_\triangleright^{-1}]))\end{aligned}$$

This isomorphism is the identity on $\sigma \Gamma$, hence we can extract the desired $\sigma_\Sigma : \sigma(\Sigma A B) \simeq \Sigma(\sigma A)((\sigma B)[\sigma_\triangleright^{-1}])$ from it.

For σ_K , note the following:

$$\begin{aligned}(\bullet \triangleright \sigma(K \Delta)) &\simeq (\sigma \bullet \triangleright \sigma(K \Delta)) \simeq \sigma(\bullet \triangleright K \Delta) \\ &\simeq \sigma \Delta \simeq (\bullet \triangleright K(\sigma \Delta))\end{aligned}$$

This yields a type isomorphism $\sigma(K \Delta) \simeq K(\sigma \Delta)$ in the empty context, and we can use the functorial action of $\epsilon : \text{Sub } \Gamma \bullet$ to weaken it to any Γ context.

For σ_{ld} , both component morphisms can be constructed by **refl** and equality reflection, and the morphisms are inverses by UIP. We omit here the verification of naturality and that σ preserves term and substitution formers in the fl-structure. \square

Identity and composition

id : **Sub** Γ Γ is defined in the obvious way, with identities for underlying functions and for preservation morphisms.

For $\sigma \circ \delta$, the underlying functions are given by function composition, and the preservation morphisms are given as follows:

$$\begin{aligned} (\sigma \circ \delta)_{\bullet}^{-1} &\equiv \sigma \delta_{\bullet}^{-1} \circ \delta_{\bullet}^{-1} \\ (\sigma \circ \delta)_{\triangleright}^{-1} &\equiv \sigma \delta_{\triangleright}^{-1} \circ \delta_{\triangleright}^{-1} \end{aligned}$$

It is easy to verify the left and right identity laws and associativity for $- \circ -$.

Lemma 6. The derived preservation isomorphisms for the fl-structure can be decomposed analogously; all derived isomorphisms in **id** are identities, and we have

$$\begin{aligned} (\sigma \circ \delta)_{\Sigma} &\equiv \sigma \delta_{\Sigma} \circ \delta_{\Sigma} \\ (\sigma \circ \delta)_{\mathbf{K}} &\equiv \sigma \delta_{\mathbf{K}} \circ \delta_{\mathbf{K}} \\ (\sigma \circ \delta)_{\text{ld}} &\equiv \sigma \delta_{\text{ld}} \circ \delta_{\text{ld}} \end{aligned}$$

On the right sides, $- \circ -$ refers to composition of type morphisms.

Proof. In the case of **ld**, the equations hold immediately by UIP. For Σ and \mathbf{K} , we prove by *flwf* computation and straightforward unfolding of definitions. \square

Empty context

The empty context \bullet : **Con** is the same as before, i.e. the terminal *flwf*. Since the unique ϵ : **Sub** Γ \bullet morphism strictly preserves all structure, it also a weak morphism.

Types

We define **Ty** Γ : **Set** as the type of split *flwf*-isofibrations over Γ . This consists of a displayed *flwf* together with *iso-cleaving* structure. For the displayed *flwf*

part, we reuse previous notation from Section 4.2.4. For the iso-cleaving, we make some auxiliary definitions first.

Definition 56 (Displayed type categories). For each $\Gamma : \mathbf{Con}_A \underline{\Gamma}$, there is a displayed category over the type category $\mathbf{Ty}_\Gamma \underline{\Gamma}$, whose objects over $\underline{A} : \mathbf{Ty}_\Gamma \underline{\Gamma}$ are elements of $\mathbf{Ty}_A \Gamma \underline{A}$, and displayed morphisms over $\underline{t} : \mathbf{Tm}_\Gamma (\underline{\Gamma} \triangleright \underline{A}) (\underline{B}[\underline{p}])$ are elements of $\mathbf{Tm}_A (\Gamma \triangleright A) (B[\underline{p}]) \underline{t}$. The identity morphism is given by \mathbf{q}_A , and the composition of t and u is $t[\mathbf{p}_A, u]$. Analogously to Definition 43, this extends to a displayed split indexed category.

Definition 57 (Displayed isomorphisms). A *displayed context isomorphism* over $\underline{\sigma} : \underline{\Gamma} \simeq \underline{\Delta}$, notated $\sigma : \Gamma \simeq_{\underline{\sigma}} \Delta$, is an invertible displayed morphism $\sigma : \mathbf{Sub}_A \Gamma \Delta \underline{\sigma}$, with inverse $\sigma^{-1} : \mathbf{Sub}_A \Delta \Gamma \underline{\sigma}^{-1}$. A *displayed type isomorphism* over $\underline{t} : \underline{A} \simeq \underline{B}$, notated $t : A \simeq_{\underline{t}} B$, is an isomorphism in a displayed type category.

Definition 58. A *vertical morphism* lies over an identity morphism. We use this definition for context morphisms (substitutions) and type morphisms as well.

Definition 59 (Context iso-cleaving). This lifts a base context isomorphism to a displayed one. It consists of

$$\begin{aligned} \text{coe} & : \underline{\Gamma} \simeq \underline{\Delta} \rightarrow \mathbf{Con}_A \underline{\Gamma} \rightarrow \mathbf{Con}_A \underline{\Delta} \\ \text{coh} & : (\underline{\sigma} : \underline{\Gamma} \simeq \underline{\Delta})(\Gamma : \mathbf{Con}_A \underline{\Gamma}) \rightarrow \Gamma \simeq_{\underline{\sigma}} \text{coe } \underline{\sigma} \Gamma \\ \text{coe}^{\text{id}} & : \text{coe id } \Gamma \equiv \Gamma \\ \text{coe}^\circ & : \text{coe } (\underline{\sigma} \circ \underline{\delta}) \Gamma \equiv \text{coe } \underline{\sigma} (\text{coe } \underline{\delta} \Gamma) \\ \text{coh}^{\text{id}} & : \text{coh id } \Gamma \equiv \text{id} \\ \text{coh}^\circ & : \text{coh } (\underline{\sigma} \circ \underline{\delta}) \Gamma \equiv \text{coh } \underline{\sigma} (\text{coh } \underline{\delta} \Gamma) \circ \text{coh } \underline{\delta} \Gamma \end{aligned}$$

Here, coe and coh abbreviate “coercion” and “coherence” respectively.

Definition 60 (Type iso-cleaving). This consists of

$$\begin{aligned} \text{coe} & : \underline{A} \simeq \underline{B} \rightarrow \mathbf{Ty}_A \Gamma \underline{A} \rightarrow \mathbf{Ty}_A \Gamma \underline{B} \\ \text{coh} & : (\underline{t} : \underline{A} \simeq \underline{B})(A : \mathbf{Ty}_A \Gamma \underline{A}) \rightarrow A \simeq_{\underline{t}} \text{coe } \underline{t} A \\ \text{coe}^{\text{id}} & : \text{coe id } A \equiv A \\ \text{coe}^\circ & : \text{coe } (\underline{t} \circ \underline{\delta}) A \equiv \text{coe } \underline{t} (\text{coe } \underline{\delta} A) \\ \text{coh}^{\text{id}} & : \text{coh id } A \equiv \text{id} \\ \text{coh}^\circ & : \text{coh } (\underline{t} \circ \underline{\delta}) A \equiv \text{coh } \underline{t} (\text{coh } \underline{\delta} A) \circ \text{coh } \underline{\delta} A \end{aligned}$$

Additionally, for $\sigma : \text{Sub}_{\mathbf{A}} \Gamma \Delta \underline{\sigma}$, we have

$$\begin{aligned} \text{coe}[] : \text{coe}(\underline{t}[\underline{\sigma}]) (A[\sigma]) &\equiv (\text{coe } \underline{t} A)[\sigma] \\ \text{coh}[] : \text{coh}(\underline{t}[\underline{\sigma} \circ \mathbf{p}, \mathbf{q}]) (A[\sigma]) &\equiv (\text{coh } \underline{t} A)[\sigma] \end{aligned}$$

Definition 61. A *split flcwf isofibration* is a displayed flCwF equipped with iso-cleaving for contexts and types.

Remark. It is not possible to model types as fibrations or opfibrations, because we have no restriction on the variance of ToS types. For example, the type which extends a pointed set signature to a natural number signature, is neither a fibration nor an opfibration.

Type substitution

We aim to define $-[-] : \mathbf{Ty} \Delta \rightarrow \mathbf{Sub} \Gamma \Delta \rightarrow \mathbf{Ty} \Gamma$, such that $\mathbf{A}[\text{id}] \equiv \mathbf{A}$ and $\mathbf{A}[\sigma \circ \delta] \equiv \mathbf{A}[\sigma][\delta]$. As before, the underlying sets are given by simple composition:

$$\begin{aligned} \text{Con}_{\mathbf{A}[\sigma]} \underline{\Gamma} &::= \text{Con}_{\mathbf{A}} (\sigma \underline{\Gamma}) \\ \text{Sub}_{\mathbf{A}[\sigma]} \Gamma \Delta \underline{\sigma} &::= \text{Sub}_{\mathbf{A}} \Gamma \Delta (\sigma \underline{\sigma}) \\ \text{Ty}_{\mathbf{A}[\sigma]} \Gamma \underline{A} &::= \text{Ty}_{\mathbf{A}} \Gamma (\sigma \underline{A}) \\ \text{Tm}_{\mathbf{A}[\sigma]} \Gamma A \underline{t} &::= \text{Tm}_{\mathbf{A}} \Gamma A (\sigma \underline{t}) \end{aligned}$$

The difference from the finitary case is that instead of preservation equations, we have isomorphisms, coercions and coherence. However, we can recover essentially the same reasoning as before, because all the previous transports still work. Context and type formers are given by coercing \mathbf{A} structures along preservation isomorphisms by σ . For example:

$$\begin{aligned} \bullet_{\mathbf{A}[\sigma]} &::= \text{coe } \sigma_{\bullet}^{-1} \bullet_{\mathbf{A}} \\ \Gamma \triangleright_{\mathbf{A}[\sigma]} A &::= \text{coe } \sigma_{\triangleright}^{-1} (\Gamma \triangleright_{\mathbf{A}} A) \\ \text{Id}_{\mathbf{A}[\sigma]} t u &::= \text{coe } \sigma_{\text{Id}}^{-1} (\text{Id}_{\mathbf{A}} t u) \\ \mathbf{K}_{\mathbf{A}[\sigma]} \Delta &::= \text{coe } \sigma_{\mathbf{K}}^{-1} (\mathbf{K}_{\mathbf{A}} \Delta) \end{aligned}$$

Term and substitution formers are given by composing **coh**-lifted isomorphisms with term and substitution formers from **A**. For example:

$$\begin{aligned} \epsilon_{\mathbf{A}[\sigma]} &: \equiv \text{coh } \sigma_{\bullet}^{-1} \bullet_{\mathbf{A}} \circ \epsilon_{\mathbf{A}} \\ \mathbf{p}_{\mathbf{A}[\sigma]} &: \equiv \mathbf{p}_{\mathbf{A}} \circ (\text{coh } \sigma_{\triangleright}^{-1} (\Gamma \triangleright A))^{-1} \\ (\sigma, \mathbf{A}[\sigma] \ t) &: \equiv \text{coh } \sigma_{\triangleright}^{-1} (\Delta \triangleright A) \circ (\sigma, \mathbf{A} \ t) \end{aligned}$$

As we mentioned, only weak **K** is supported in **A**[σ]. For strict **K** we would have to show:

$$\text{Sub}_{\mathbf{A}} \Gamma \Delta (\sigma \ \underline{\sigma}) \equiv \text{Tm}_{\mathbf{A}} \Gamma (\text{coe } \sigma_{\mathbf{K}}^{-1} (\mathbf{K}_{\mathbf{A}} \Delta)) (\sigma \ \underline{\sigma})$$

By strict **K** in **A**, it would be enough to show

$$\text{Tm}_{\mathbf{A}} \Gamma (\mathbf{K}_{\mathbf{A}} \Delta) (\sigma \ \underline{\sigma}) \equiv \text{Tm}_{\mathbf{A}} \Gamma (\text{coe } \sigma_{\mathbf{K}}^{-1} (\mathbf{K}_{\mathbf{A}} \Delta)) (\sigma \ \underline{\sigma})$$

But there is no reason why these sets should be equal, so we instead produce an isomorphism.

Equations for term and type substitution follow from naturality of preservation isomorphisms in σ , $\text{coe}[]$, $\text{coh}[]$ and substitution equations in **A**.

Iso-cleaving is given by iso-cleaving in **A** and the action of σ on isomorphisms, so that we have $\text{coe}_{\mathbf{A}[\sigma]} \underline{\sigma} \Gamma \equiv \text{coe}_{\mathbf{A}} (\sigma \ \underline{\sigma}) \Gamma$ and $\text{coh}_{\mathbf{A}[\sigma]} \underline{\sigma} \Gamma \equiv \text{coh}_{\mathbf{A}} (\sigma \ \underline{\sigma}) \Gamma$.

Functoriality of type substitution, i.e. $\mathbf{A}[\text{id}] \equiv \mathbf{A}$ and $\mathbf{A}[\sigma \circ \delta] \equiv \mathbf{A}[\sigma][\delta]$, follows from Lemma 6 and split cleaving given by coe^{id} , coe° , coh^{id} and coh° laws in **A**.

Terms

$\mathbf{Tm} \Gamma \mathbf{A} : \text{Set}$ is defined as the type of *weak flCwF sections* of **A**. The underlying functions of $t : \mathbf{Tm} \Gamma \mathbf{A}$ are as follows:

$$\begin{aligned} t &: (\underline{\Gamma} : \text{Con}_{\Gamma}) \rightarrow \text{Con}_{\mathbf{A}} \underline{\Gamma} \\ t &: (\underline{\sigma} : \text{Sub}_{\Gamma} \underline{\Gamma} \underline{\Delta}) \rightarrow \text{Sub}_{\mathbf{A}} (t \ \underline{\Gamma}) (t \ \underline{\Delta}) \underline{\sigma} \\ t &: (\underline{A} : \text{Ty}_{\Gamma}) \rightarrow \text{Ty}_{\mathbf{A}} (t \ \underline{\Gamma}) \underline{A} \\ t &: (t : \mathbf{Tm}_{\Gamma} \underline{\Gamma} \underline{A}) \rightarrow \mathbf{Tm}_{\mathbf{A}} (t \ \underline{\Gamma}) (t \ \underline{A}) \underline{t} \end{aligned}$$

Such that

$$1. \ t (\underline{A}[\sigma]) \equiv (t \ \underline{A}) [t \ \underline{\sigma}]$$

2. $\mathbf{t}(\mathbf{t}[\underline{\sigma}]) \equiv (\mathbf{t}\mathbf{t})[\mathbf{t}\underline{\sigma}]$
3. The unique $\epsilon_A : \mathbf{Sub}(\mathbf{t}\bullet) \bullet \mathbf{id}$ has a vertical retraction.
4. Each $(\mathbf{t}\mathbf{p}, \mathbf{t}\mathbf{q}) : \mathbf{Sub}(\mathbf{t}(\underline{\Gamma} \triangleright \underline{A}))(\mathbf{t}\underline{\Gamma} \triangleright \mathbf{t}\underline{A}) \mathbf{id}$ has a vertical inverse.

Similarly to what we had in **Sub**, we denote the evident preservation isomorphisms as $\mathbf{t}_\bullet : \mathbf{t}\bullet \simeq_{\mathbf{id}} \bullet$ and $\mathbf{t}_\triangleright : \mathbf{t}(\underline{\Gamma} \triangleright \underline{A}) \simeq_{\mathbf{id}} \mathbf{t}\underline{\Gamma} \triangleright \mathbf{t}\underline{A}$. In short, weak sections are dependently typed analogues of weak morphisms, with dependent underlying functions and displayed preservation isomorphisms. We also have the derived fl-preservation isomorphisms.

Theorem 7. *A weak section $\mathbf{t} : \mathbf{Tm} \Gamma \mathbf{A}$ preserves fl-structure up to vertical type isomorphisms, that is, the following are derivable:*

$$\begin{aligned} \mathbf{t}_\Sigma : \mathbf{t}(\Sigma \underline{A} \underline{B}) &\simeq_{\mathbf{id}} \Sigma(\mathbf{t}\underline{A})((\mathbf{t}\underline{B})[\mathbf{t}_\triangleright^{-1}]) \\ \mathbf{t}_K : \mathbf{t}(K \underline{\Delta}) &\simeq_{\mathbf{id}} K(\mathbf{t}\underline{\Delta}) \\ \mathbf{t}_{\mathbf{id}} : \mathbf{t}(\mathbf{Id} \mathbf{t} \underline{u}) &\simeq_{\mathbf{id}} \mathbf{Id}(\mathbf{t}\mathbf{t})(\mathbf{t}\underline{u}) \end{aligned}$$

Also, the above isomorphisms are natural in the sense of Theorem 6, and \mathbf{t} preserves term and substitution formers in the fl-structure.

Proof. The construction of isomorphisms is the same as in Theorem 6. Indeed, every construction there has a displayed counterpart which we can use here. \square

We note though that the move from Theorem 6 to here is not simply a logical predicate translation, because we are only lifting the codomain of a weak morphism to a displayed version, and we leave the domain non-displayed. We leave to future work the investigation of such asymmetrical (or “modal”) logical predicate translations.

Term substitution

$-[-] : \mathbf{Tm} \Delta \mathbf{A} \rightarrow (\sigma : \mathbf{Sub} \Gamma \Delta) \rightarrow \mathbf{Tm} \Gamma(\mathbf{A}[\sigma])$ is given similarly to $- \circ -$ in Section 5.2.2. Underlying functions are given by function composition, and preservation morphisms are also similar:

$$\begin{aligned} (\mathbf{t}[\sigma])_\bullet^{-1} &\equiv \mathbf{t}\sigma_\bullet^{-1} \circ \mathbf{t}_\bullet^{-1} \\ (\mathbf{t}[\sigma])_\triangleright^{-1} &\equiv \mathbf{t}\sigma_\triangleright^{-1} \circ \mathbf{t}_\triangleright^{-1} \end{aligned}$$

We also have the same decomposition of derived isomorphisms as in Lemma 6. We do not have to show functoriality of term substitution here, since that is derivable in any cwf, see e.g. [KKA19].

Comprehension

$\Gamma \triangleright \mathbf{A} : \mathbf{Con}$ is defined as the total flcwf of \mathbf{A} , in exactly the same way as in the finitary case, since the additional iso-cleaving structure plays no role in the result. $\mathbf{p} : \mathbf{Sub}(\Gamma \triangleright \mathbf{A}) \Gamma$ and $\mathbf{q} : \mathbf{Tm}(\Gamma \triangleright \mathbf{A})(\mathbf{A}[\mathbf{p}])$ are likewise unchanged; they are strict morphisms, so also automatically weak morphisms. Substitution extension (σ, t) is given by pointwise combining σ and t , e.g. $\mathbf{Con}_{(\sigma, t)} \underline{\Gamma} \equiv (\sigma \underline{\Gamma}, t \underline{\Gamma})$.

Strict constant families

We have the same definition for $\mathbf{K} \Delta : \mathbf{T} \mathbf{y} \Gamma$ as in the finitary case, although we need to define iso-cleaving in addition. Fortunately, coercions and coherences are all trivial, because $\mathbf{K} \Delta$ does not actually depend on Γ .

$$\begin{aligned} \mathbf{coe}_{\mathbf{K} \Delta} \underline{\sigma} \Gamma &\equiv \Gamma \\ \mathbf{coe}_{\mathbf{K} \Delta} \underline{t} A &\equiv A \end{aligned}$$

Universe

$\mathbf{U} : \mathbf{T} \mathbf{y} \Gamma$ is exactly the same as before. We define it as the type which is constantly the flcwf of inner types, so it inherits the trivial iso-cleaving from \mathbf{K} .

$\mathbf{El} \mathbf{a} : \mathbf{T} \mathbf{y} \Gamma$ is again the displayed flcwf of the elements of $\mathbf{a} : \mathbf{Tm} \Gamma \mathbf{U}$. The underlying sets are unchanged:

$$\begin{aligned} \mathbf{Con}_{\mathbf{El} \mathbf{a}} \underline{\Gamma} &\equiv \mathbf{Tm}_0(\mathbf{a} \underline{\Gamma}) \\ \mathbf{Sub}_{\mathbf{El} \mathbf{a}} \Gamma \Delta \underline{\sigma} &\equiv \mathbf{a} \underline{\sigma} \Gamma \equiv \Delta \\ \mathbf{T} \mathbf{y}_{\mathbf{El} \mathbf{a}} \Gamma \underline{A} &\equiv \mathbf{Tm}_0(\mathbf{a} \underline{A} \Gamma) \\ \mathbf{Tm}_{\mathbf{El} \mathbf{a}} \Gamma A \underline{t} &\equiv \mathbf{a} \underline{t} \Gamma \equiv A \end{aligned}$$

We need to adjust definitions to show that $\mathbf{El} \mathbf{a}$ supports all required structure. Previously, all context and type formers were inherited from \mathbf{U} , since \mathbf{a} strictly preserved them. Now, \mathbf{a} preserves structure up to (definitional) isomorphism of inner types. Hence, the adjustments are quite mechanical; they are like wrapping

all definitions in “unary record constructors” given by preservation isomorphisms. For example:

$$\begin{aligned} \bullet \mathbf{El} \, a & \quad \equiv \, \mathbf{a}_\bullet^{-1} \, \mathbf{tt} \\ (\Gamma \triangleright_{\mathbf{El} \, a} A) & \quad \equiv \, \mathbf{a}_{\triangleright}^{-1} (\Gamma, A) \end{aligned}$$

We likewise use preservation isomorphisms to define \mathbf{K} , \mathbf{Id} and Σ . Context coercion is $\mathbf{coe} \, \underline{\sigma} \, \Gamma \equiv \mathbf{a} \, \underline{\sigma} \, \Gamma$. Type coercion, for $A : \mathbf{a} \, \underline{A} \, \Gamma$ is given as $\mathbf{coe} \, \underline{t} \, A \equiv \mathbf{a} \, \underline{t} \, (\mathbf{a}_{\triangleright}^{-1} (\Gamma, A))$.

Unit type

$\top : \mathbf{Tm} \, \Gamma \, \mathbf{U}$ is the constantly \top_0 morphism, i.e. it maps objects to \top_0 and types to $\lambda _ . \top_0$, and maps morphisms and terms to the identity function. It clearly preserves \bullet and $-\triangleright-$ up to isomorphism.

Sigma type

For $\mathbf{a} : \mathbf{Tm} \, \Gamma \, \mathbf{U}$ and $\mathbf{b} : \mathbf{Tm} (\Gamma \triangleright \mathbf{El} \, \mathbf{a}) \, \mathbf{U}$, we define $\Sigma \, \mathbf{a} \, \mathbf{b} : \mathbf{Tm} \, \Gamma \, \mathbf{U}$ as the component-wise Σ of \mathbf{a} and \mathbf{b} . For the action on $\underline{\Gamma} : \mathbf{Con}_\Gamma$, we have:

$$\begin{aligned} (\Sigma \, \mathbf{a} \, \mathbf{b}) \, \underline{\Gamma} & : \top_{y_0} \\ (\Sigma \, \mathbf{a} \, \mathbf{b}) \, \underline{\Gamma} & \equiv (\alpha : \mathbf{a} \, \underline{\Gamma}) \times \mathbf{b} (\underline{\Gamma}, \alpha) \end{aligned}$$

For the action on $\underline{\sigma} : \mathbf{Sub} \, \underline{\Gamma} \, \underline{\Delta}$, we have:

$$\begin{aligned} (\Sigma \, \mathbf{a} \, \mathbf{b}) \, \underline{\sigma} & : (\alpha : \mathbf{a} \, \underline{\Gamma}) \times \mathbf{b} (\underline{\Gamma}, \alpha) \rightarrow (\alpha : \mathbf{a} \, \underline{\Delta}) \times \mathbf{b} (\underline{\Delta}, \alpha) \\ (\Sigma \, \mathbf{a} \, \mathbf{b}) \, \underline{\sigma} & \equiv \lambda (\alpha, \beta). (\mathbf{a} \, \underline{\sigma} \, \alpha, \mathbf{b} (\underline{\sigma}, \mathbf{refl}) \, \beta) \end{aligned}$$

Above, the second field should have type $\mathbf{b} (\underline{\Delta}, \mathbf{a} \, \underline{\sigma} \, \alpha)$, while $\beta : \mathbf{b} (\underline{\Gamma}, \alpha)$. Therefore we need a morphism in $\Gamma \triangleright \mathbf{El} \, \mathbf{a}$ from $(\underline{\Gamma}, \alpha)$ to $(\underline{\Delta}, \mathbf{a} \, \underline{\sigma} \, \alpha)$, which is defined as $(\underline{\sigma}, \mathbf{refl})$, where $\mathbf{refl} : \mathbf{a} \, \underline{\sigma} \, \alpha \equiv \mathbf{a} \, \underline{\sigma} \, \alpha$. The action on $\underline{A} : \mathbf{Ty} \, \underline{\Gamma}$ is

$$\begin{aligned} (\Sigma \, \mathbf{a} \, \mathbf{b}) \, \underline{A} & : (\alpha : \mathbf{a} \, \underline{\Gamma}) \times \mathbf{b} (\underline{\Gamma}, \alpha) \rightarrow \top_{y_0} \\ (\Sigma \, \mathbf{a} \, \mathbf{b}) \, \underline{A} & \equiv \lambda (\alpha, \beta). (\alpha' : \mathbf{a} \, \underline{A} \, \alpha) \times \mathbf{b} (\underline{A}, \alpha') \, \beta \end{aligned}$$

Here we are somewhat running out of notation: we use α' to refer to a *type* over $\alpha : \mathbf{a} \, \underline{\Gamma}$ in the displayed cwf of elements $\mathbf{El} \, \mathbf{a}$. The action on terms is analogous:

$$\begin{aligned} (\Sigma \, \mathbf{a} \, \mathbf{b}) \, \underline{t} & : ((\alpha, \beta) : (\alpha : \mathbf{a} \, \underline{\Gamma}) \times \mathbf{b} (\underline{\Gamma}, \alpha)) \rightarrow (\alpha' : \mathbf{a} \, \underline{A} \, \alpha) \times \mathbf{b} (\underline{A}, \alpha') \, \beta \\ (\Sigma \, \mathbf{a} \, \mathbf{b}) \, \underline{t} & \equiv \lambda (\alpha, \beta). (\mathbf{a} \, \underline{t} \, \alpha, \mathbf{b} (\underline{t}, \mathbf{refl}) \, \beta) \end{aligned}$$

For the preservation of \bullet , we need to show $(\Sigma \mathbf{a} \mathbf{b}) \bullet \simeq \top_0$. Unfolding the definition, we get $((\alpha : \mathbf{a} \bullet) \times \mathbf{b}(\bullet, \alpha)) \simeq \top_0$. This holds since $\mathbf{a} \bullet \simeq \top_0$, so $\mathbf{a} \bullet$ is contractible, thus $(\bullet, \alpha) \equiv \bullet_{\Gamma \triangleright \mathbf{El} \mathbf{a}}$, and we also know $\mathbf{b} \bullet \simeq \top_0$. For the preservation $-\triangleright-$, we need

$$(\Sigma \mathbf{a} \mathbf{b}) (\underline{\Gamma} \triangleright \underline{A}) \simeq (\gamma : (\Sigma \mathbf{a} \mathbf{b}) \underline{\Gamma}) \times (\Sigma \mathbf{a} \mathbf{b}) \underline{A} \gamma$$

Unfolding definitions and reassociating Σ on the right side:

$$\begin{aligned} & (\alpha : \mathbf{a} (\underline{\Gamma} \triangleright \underline{A})) \times \mathbf{b}((\underline{\Gamma} \triangleright \underline{A}), \alpha) \\ & \simeq \\ & (\alpha : \mathbf{a} \underline{\Gamma}) \times (\beta : \mathbf{b}(\underline{\Gamma}, \alpha)) \times (\alpha' : \mathbf{a} \underline{A} \alpha) \times \mathbf{b}(\underline{A}, \alpha') \beta \end{aligned}$$

Since $\mathbf{a}_{\triangleright} : \mathbf{a} (\underline{\Gamma} \triangleright \underline{A}) \simeq (\alpha : \mathbf{a} \underline{\Gamma}) \times (\beta : \mathbf{b}(\underline{\Gamma}, \alpha))$, we can rewrite the left side using pattern matching notation as

$$(\mathbf{a}_{\triangleright}^{-1}(\gamma, \alpha) : \mathbf{a} (\underline{\Gamma} \triangleright \underline{A})) \times \mathbf{b}((\underline{\Gamma} \triangleright \underline{A}), (\gamma, \alpha))$$

Now, since $((\underline{\Gamma} \triangleright \underline{A}), (\gamma, \alpha)) \equiv (\underline{\Gamma}, \gamma)_{\triangleright \Gamma \triangleright \mathbf{El} \mathbf{a}}(\underline{A}, \alpha)$, we know that $\mathbf{b}((\underline{\Gamma} \triangleright \underline{A}), (\gamma, \alpha))$ is also isomorphic to the evident Σ type, and the preservation isomorphism follows.

Projections and pairing for $\Sigma \mathbf{a} \mathbf{b}$ are defined in the obvious way by component-wise projection and pairing.

Identity

For \mathbf{t} and \mathbf{u} in $\mathbf{Tm} \Gamma (\mathbf{El} \mathbf{a})$, we define $\mathbf{ld} \mathbf{t} \mathbf{u} : \mathbf{Tm} \Gamma \mathbf{U}$ as expressing pointwise equality of weak sections. We rely on the assumption that \mathbf{Ty}_0 has identity type.

$$\begin{aligned} (\mathbf{ld} \mathbf{t} \mathbf{u}) \underline{\Gamma} &::= (\mathbf{t} \underline{\Gamma} = \mathbf{u} \underline{\Gamma}) \\ (\mathbf{ld} \mathbf{t} \mathbf{u}) \underline{A} &::= \lambda e. (\mathbf{t} \underline{A} = \mathbf{u} \underline{A}) \end{aligned}$$

Above, $\mathbf{t} \underline{A} = \mathbf{u} \underline{A}$ is well-typed because of $e : \mathbf{t} \underline{\Gamma} = \mathbf{u} \underline{\Gamma}$. For substitutions, we have to complete a square of equalities:

$$(\mathbf{ld} \mathbf{t} \mathbf{u}) (\underline{\sigma} : \mathbf{Sub} \underline{\Gamma} \underline{\Delta}) : (\mathbf{t} \underline{\Gamma} = \mathbf{u} \underline{\Gamma}) \rightarrow (\mathbf{t} \underline{\Delta} = \mathbf{u} \underline{\Delta})$$

This can be given by $\mathbf{t} \underline{\sigma} : \mathbf{a} \underline{\sigma} (\mathbf{t} \underline{\Gamma}) = \mathbf{t} \underline{\Delta}$ and $\mathbf{u} \underline{\sigma} : \mathbf{a} \underline{\sigma} (\mathbf{u} \underline{\Gamma}) = \mathbf{u} \underline{\Delta}$. The action on terms is analogous.

The \bullet -preservation $(\mathbf{t} \bullet = \mathbf{u} \bullet) \simeq \top_0$ follows from $\mathbf{a} \bullet \simeq \top_0$. For \triangleright -preservation, we need to show

$$(\mathbf{t} (\Gamma \triangleright \underline{A}) = \mathbf{u} (\Gamma \triangleright \underline{A})) \simeq ((e : \mathbf{t} \Gamma = \mathbf{u} \Gamma) \times (\mathbf{t} \underline{A} = \mathbf{u} \underline{A}))$$

This follows from \triangleright -preservation by \mathbf{a} . Equality reflection and **refl** : **Id** $\mathbf{t} \mathbf{t}$ are also evident.

Infinitary function type

For $Ix : \mathsf{Ty}_0$ and $\mathbf{b} : Ix \rightarrow \mathbf{Tm} \Gamma \mathbf{U}$, we aim to define $\Pi^{\text{inf}} Ix \mathbf{b} : \mathbf{Tm} \Gamma \mathbf{U}$. The underlying functions are:

$$\begin{aligned} (\Pi^{\text{inf}} Ix \mathbf{b}) \Gamma &:= (i : Ix) \rightarrow \mathbf{b} i \Gamma \\ (\Pi^{\text{inf}} Ix \mathbf{b}) \underline{\sigma} &:= \lambda i. \mathbf{b} i \underline{\sigma} \\ (\Pi^{\text{inf}} Ix \mathbf{b}) \underline{A} &:= \lambda \Gamma. (i : Ix) \rightarrow \mathbf{b} i \underline{A} (\Gamma a) \\ (\Pi^{\text{inf}} Ix \mathbf{b}) \underline{t} &:= \lambda i. \mathbf{b} i \underline{t} \end{aligned}$$

We rely on Π in the inner theory. The preservation isomorphisms are pointwise inherited from \mathbf{b} . One direction of the isomorphisms is defined as follows. Note that $\bullet_{\mathbf{U}} \equiv \top$ and $\triangleright_{\mathbf{U}}$ is Σ .

$$\begin{aligned} (\Pi^{\text{inf}} Ix \mathbf{b})_{\bullet}^{-1} &: \top \rightarrow (\Pi^{\text{inf}} Ix \mathbf{b}) \bullet \\ (\Pi^{\text{inf}} Ix \mathbf{b})_{\bullet}^{-1} &:= \lambda _ i. (\mathbf{b} i)_{\bullet}^{-1} \text{tt} \\ (\Pi^{\text{inf}} Ix \mathbf{b})_{\triangleright}^{-1} &: (\Gamma : (\Pi^{\text{inf}} Ix \mathbf{b}) \Gamma) \times ((\Pi^{\text{inf}} Ix \mathbf{b}) \underline{A} \Gamma) \\ &\rightarrow (\Pi^{\text{inf}} Ix \mathbf{b}) (\Gamma \triangleright \underline{A}) \\ (\Pi^{\text{inf}} Ix \mathbf{b})_{\triangleright}^{-1} &:= \lambda (\Gamma, A) i. (\mathbf{b} i)_{\triangleright}^{-1} (\Gamma i, A i) \end{aligned}$$

Inductive function type

For $\mathbf{a} : \mathbf{Tm} \Gamma \mathbf{U}$ and $\mathbf{B} : \mathbf{Ty} (\Gamma \triangleright \mathbf{El} \mathbf{a})$, we aim to define $\Pi \mathbf{a} \mathbf{B} : \mathbf{Ty} \Gamma$. The underlying sets are unchanged.

$$\begin{aligned} \text{Con}_{\Pi \mathbf{a} \mathbf{B}} \Gamma &:= (\gamma : \mathbf{a} \Gamma) \rightarrow \text{Con}_{\mathbf{B}} (\Gamma, \gamma) \\ \text{Sub}_{\Pi \mathbf{a} \mathbf{B}} \Gamma \Delta \underline{\sigma} &:= (\gamma : \mathbf{a} \Gamma) \rightarrow \text{Sub}_{\mathbf{B}} (\Gamma \gamma) (\Delta (\mathbf{a} \underline{\sigma} \gamma)) (\underline{\sigma}, \text{refl}) \\ \text{Ty}_{\Pi \mathbf{a} \mathbf{B}} \Gamma \underline{A} &:= \{\gamma : \mathbf{a} \Gamma\} (\alpha : \mathbf{a} \underline{A} \gamma) \rightarrow \text{Ty}_{\mathbf{B}} (\Gamma \gamma) (\underline{A}, \alpha) \\ \text{Tm}_{\Pi \mathbf{a} \mathbf{B}} \Gamma A \underline{t} &:= (\gamma : \mathbf{a} \Gamma) \rightarrow \text{Tm}_{\mathbf{B}} (\Gamma \gamma) (\mathbf{A} (\mathbf{a} \underline{t} \gamma)) (\underline{t}, \text{refl}) \end{aligned}$$

Likewise, all structure is defined pointwise using \mathbf{B} structure. Similarly to the \mathbf{EI} case, we have to sometimes fall through the defining isomorphisms for \mathbf{a} structure. For comparison, in the finitary case we had the following definition:

$$(\Gamma \triangleright_{\Pi \mathbf{a} \mathbf{B}} A) (\gamma, \alpha) := (\Gamma \gamma \triangleright_{\mathbf{B}} A \alpha)$$

Here, $(\gamma, \alpha) : \mathbf{a} (\underline{\Gamma} \triangleright \underline{A})$, so also $(\gamma, \alpha) : (\gamma : \mathbf{a} \underline{\Gamma}) \times \mathbf{a} \underline{A} \gamma$, so the Σ pattern-matching notation is justified in the definition. In the current infinitary case, we have $(\mathbf{a}_{\triangleright}, \mathbf{a}_{\triangleright}^{-1}) : \mathbf{a} (\underline{\Gamma} \triangleright \underline{A}) \simeq ((\gamma : \mathbf{a} \underline{\Gamma}) \times \mathbf{a} \underline{A} \gamma)$ instead. But we can use the intuition that set isomorphisms are like unary record types, so we can still give a pattern-matching definition:

$$(\Gamma \triangleright_{\Pi \mathbf{a} \mathbf{B}} A) (\mathbf{a}_{\triangleright}^{-1}(\gamma, \alpha)) := (\Gamma \gamma \triangleright_{\mathbf{B}} A \alpha)$$

For the definitions of other type and term formers, we likewise insert the isomorphisms appropriately. It remains to define iso-cleaving Π . Coercion is given by mapping indices backwards in $\mathbf{EI} \mathbf{a}$ and coercing outputs forwards in \mathbf{B} .

$$\begin{aligned} \text{coe } \underline{\sigma} \Gamma &:= \lambda \gamma. \text{coe}_{\mathbf{B}} (\underline{\sigma}, \text{refl}) (\Gamma (\mathbf{a} (\underline{\sigma}^{-1}) \gamma)) \\ \text{coe } \underline{t} A &:= \lambda \gamma a. \text{coe}_{\mathbf{B}} (\underline{t}, \text{refl}) (A (\mathbf{a} (\underline{t}^{-1}) (\mathbf{a}_{\triangleright}^{-1}(\gamma, a)))) \end{aligned}$$

Likewise, coh -s are given by backwards-forwards coh -s. As before, \mathbf{app} and \mathbf{lam} are defined as currying and uncurrying the underlying functions.

External function type

For $Ix : \mathbf{Set}_j$ and $\mathbf{B} : Ix \rightarrow \mathbf{Ty} \Gamma$, we define $\Pi^{\text{ext}} Ix \mathbf{B} : \mathbf{Ty} \Gamma$ as the Ix -indexed direct product of \mathbf{B} . Since the indexing is given by a metatheoretic function, every component is given in the evident pointwise way, including iso-cleaving.

5.3 Discussion of Semantics

Iso-fibrancy as a structure identity principle

The flwfs of algebras that we get from the infinitary semantics are exactly the same as in the finitary case. However, signature types are a bit more interesting. The iso-fibrancy of types can be understood as a weaker version of the *structure identity principle* in homotopy type theory.

The structure identity principle says that isomorphism of algebras is equivalent to equality of algebras. This is the same as saying that categories of algebras are univalent [AKS15]. Assuming a signature Γ and algebras $\gamma \simeq \gamma'$, we have $\gamma = \gamma'$. This equality is respected by every construction in HoTT, which implies that for any HoTT type family $F : \Gamma^A \rightarrow \mathbf{Type}$, we have a function $F \gamma \rightarrow F \gamma'$.

We get a similar but weaker statement from the infinitary semantics: for $\sigma : \gamma \simeq \gamma'$ and some ToS type $A : \mathbf{Ty} \Gamma$, we have a function $\mathbf{coe} \sigma : A^A \gamma \rightarrow A^A \gamma'$. We also have $\mathbf{coh} \sigma \alpha : \alpha \simeq_{\sigma} \mathbf{coe} \sigma \alpha$ for some $\alpha : A^A \gamma$. So we can transport over isomorphisms, but not all metatheoretic families can be transported, only those which arise as ToS types.

Of course, we can transport over multiple types, or telescopes of types too, by iterated transport. For instance, given $A : \mathbf{Ty} \Gamma$, $B : \mathbf{Ty} (\Gamma \triangleright A)$, $\alpha : A^A \gamma$ and $\beta : B^A (\gamma, \alpha)$, we can transport α first, then transport β by $(\sigma, \mathbf{coh} \sigma \alpha)$. Alternatively, we could add large Σ types to ToS, as $\Sigma : (A : \mathbf{Ty} \Gamma) \rightarrow \mathbf{Ty} (\Gamma \triangleright A) \rightarrow \mathbf{Ty} \Gamma$, thereby making iterated transport superfluous.

Variations of semantics

First, unlike in the finitary case, we have no opportunity to minimize assumptions on the inner theory. Already when we compute algebras, we need inner Π for infinitary functions, inner \top for \top , inner Σ for Σ and inner $- = -$ for \mathbf{Id} .

Second: can we add the “large” equality type, which includes sort equations, back to infinitary signatures? We dropped sort equations in this chapter because they are clearly not isofibrant. We can add them back into the mix though, at the price of dropping things from the semantics of signatures. The reason for having isofibrant types is that type formers in \mathbf{U} preserve \bullet and $-\triangleright-$ only up to isomorphism. If we drop all semantic components which depend on \bullet and $-\triangleright-$, we can drop isofibrancy too from the model, and everything works. In this case, we still get a category of algebras, plus a notion of induction, but we cannot show that initiality is equivalent to induction, as the proof of Theorem 1 depends on $-\triangleright-$. We cannot do term algebra constructions (see in Section 5.4) and lambda-encodings (in Section ??) either, since they rely on isofibrancy.

Substitutions

Add section on signature semantics of signatures!!

5.4 Term Algebras

We adapt now the previous term algebra construction to the infinitary case. We again switch to the ETT setup with cumulative universes. We assume Section 4.4.1 without any change. Also, we adapt 4.4.2 to infinitary signatures and semantics. All definitions are the same, the only change is that the $\mathbf{M}_{i,j}$ model is now the isofibrant flwf model, and we have the infinitary ToS.

5.4.1 Term Algebra Construction

The term algebra construction changes significantly. The reason is the following. In the finitary case, the key property was that “small types evaluated in the term model are sets of terms”. Formally, we had for $a : \mathbf{Tm} \Omega \mathbf{U}$ that $a^A (\Omega^T \text{id}) \equiv \mathbf{Tm} \Omega (\text{El } a)$. This is now weakened to an isomorphism, i.e. $a^A (\Omega^T \text{id}) \simeq \mathbf{Tm} \Omega (\text{El } a)$.

This is again necessary because of the closure of \mathbf{U} under type formers. For example, $\top^A (\Omega^T \text{id}) \equiv \top$, and $\mathbf{Tm} \Omega (\text{El } \top)$ is merely isomorphic to \top . We assume $\Omega : \text{Sig}_j$ for some j level, and define $-^T$ by induction on syn_j .

$$\begin{aligned} -^T : (\Gamma : \text{Con}) \quad (\nu : \text{Sub } \Omega \Gamma) &\rightarrow \Gamma^A \\ -^T : (\sigma : \text{Sub } \Gamma \Delta) (\nu : \text{Sub } \Omega \Gamma) &\rightarrow \Delta^T (\sigma \circ \nu) \simeq \sigma^A (\Gamma^T \nu) \\ -^T : (A : \text{Ty } \Gamma) \quad (\nu : \text{Sub } \Omega \Gamma) &\rightarrow \mathbf{Tm} \Omega (A[\nu]) \rightarrow A^A (\Gamma^T \nu) \\ -^T : (t : \mathbf{Tm} \Gamma A) \quad (\nu : \text{Sub } \Omega \Gamma) &\rightarrow A^T \nu (t[\nu]) \simeq_{\text{id}} t^A (\Gamma^T \nu) \end{aligned}$$

In short, interpretations of substitutions and terms are weakened to isomorphisms. By \simeq_{id} we mean a displayed isomorphism of objects in the semantic A (which is an flwf isofibration); recall Definition 57. The isomorphism is “vertical” since it lies over id .

The interpretation of the cwf is the same as before, but like in the isofibrant semantics, we have to use explicit **coe** instead of silently transporting over equalities. In the interpretations of substitutions and terms, we have to explicitly compose isomorphisms and sometimes lift them using **coh**. We give some examples. The interpretation of context formers is the same as before:

$$\begin{aligned} \bullet^T \nu &:\equiv \text{tt} \\ (\Gamma \triangleright A)^T (\nu, t) &:\equiv (\Gamma^T \nu, A^T \nu t) \end{aligned}$$

Type substitution with $\sigma : \mathbf{Sub} \Gamma \Delta$ is interpreted as coercion:

$$\begin{aligned} (A[\sigma])^T : (\nu : \mathbf{Sub} \Omega \Gamma)(t : \mathbf{Tm} \Omega (A[\sigma][\nu]) &\rightarrow A^A (\sigma^A (\Gamma^T \nu)) \\ (A[\sigma])^T \nu t &\equiv \mathbf{coe} (\sigma^T \nu) (A^T (\sigma \circ \nu) t) \end{aligned}$$

Composition of $\sigma : \mathbf{Sub} \Delta \Xi$ and $\delta : \mathbf{Sub} \Gamma \Delta$ is the following:

$$\begin{aligned} (\sigma \circ \delta)^T : (\nu : \mathbf{Sub} \Omega \Gamma) &\rightarrow \Delta^T (\sigma \circ \delta \circ \nu) \simeq \sigma^A (\delta^A (\Gamma^T \nu)) \\ (\sigma \circ \delta)^T \nu &\equiv \sigma^M (\delta^T \nu) \circ \sigma^T (\delta \circ \nu) \end{aligned}$$

Above, we have

$$\begin{aligned} \delta^T \nu &: \Xi^T (\delta \circ \nu) \simeq \delta^A (\Gamma^T \nu) \\ \sigma^M (\delta^T \nu) &: \sigma^A (\Xi^T (\delta \circ \nu)) \simeq \sigma^A (\delta^A (\Gamma^T \nu)) \\ \sigma^T (\delta \circ \nu) &: \Delta^T (\sigma \circ \delta \circ \nu) \simeq \sigma^A (\Xi^T (\delta \circ \nu)) \end{aligned}$$

Hence, the type of the composition in the definition checks out. We make use of the fact that σ^M sends an isomorphism in Γ to an isomorphism in Δ .

Substitution extension is a somewhat more complicated case. We want to interpret the extension of $\sigma : \mathbf{Sub} \Gamma \Delta$ with $t : \mathbf{Tm} \Gamma (A[\sigma])$:

$$(\sigma, t)^T : (\nu : \mathbf{Sub} \Omega \Gamma) \rightarrow (\Delta \triangleright A)^T ((\sigma, t) \circ \nu) \simeq (\sigma, t)^A (\Gamma^T \nu)$$

The goal is an isomorphism in the semantic $\Gamma \triangleright A$ category, i.e. the total category of A . Every isomorphism in $\Gamma \triangleright A$ arises as packing together a Γ isomorphism and a displayed A isomorphism over it. We can compute the type further:

$$(\sigma, t)^T : (\nu : \mathbf{Sub} \Omega \Gamma) \rightarrow (\Delta^T (\sigma \circ \nu), A^T (\sigma \circ \nu) (t[\nu])) \simeq (\sigma^A (\Gamma^T \nu), t^A (\Gamma^T \nu))$$

We can exhibit $\sigma^T \nu : \Delta^T (\sigma \circ \nu) \simeq \sigma^A (\Gamma^T \nu)$ as the base component of the goal isomorphism. Now we need a displayed isomorphism over it. Following the pattern, we may try $t^T \nu$:

$$t^T \nu : (A[\sigma])^T \nu (t[\nu]) \simeq_{\text{id}} t^A (\Gamma^T \nu)$$

Computing the type:

$$t^T \nu : \mathbf{coe} (\sigma^T \nu) (A^T (\sigma \circ \nu) (t[\nu])) \simeq_{\text{id}} t^A (\Gamma^T \nu)$$

So this is not quite what is needed; we want a displayed iso over $\sigma^T \nu$, but we have something over id . We can fix this using \mathbf{coh} :

$$\mathbf{coh} (\sigma^T \nu) (A^T (\sigma \circ \nu) (t[\nu])) : A^T (\sigma \circ \nu) (t[\nu]) \simeq_{\sigma^T \nu} \mathbf{coe} (\sigma^T \nu) (A^T (\sigma \circ \nu) (t[\nu]))$$

The composition of $t^T \nu$ and the above now checks out:

$$(\sigma, t)^T \nu \equiv (\sigma^T \nu, t^T \nu \circ \mathbf{coh}(\sigma^T \nu)(A^T(\sigma \circ \nu)(t[\nu])))$$

We omit the rest of the cwf interpretation. It should be apparent that explicit **coe** and **coh**-handling is fairly technical. It should be noted though that in a proof assistant, the finitary and infinitary constructions would be of similar difficulty, because there we cannot rely on equality reflection and implicit transports to magically tidy up the formalization. In fact, even in the finitary case it would be a good idea to structure the formalization around coercions and coherences.

The high-level explanation for why the weakened constructions continue to work, is the same as what we gave in the section on iso-fibrant semantics: we do nothing which would violate stability under isomorphisms; additionally, because our isofibrations are *split*, coercion and coherence compute strictly on identities and compositions, which ensures that conversion equations in the syntax are respected. For example, functoriality of the type substitution defined above relies on **coe** computation on identity and composition.

Universe

The universe is interpreted as follows.

$$\mathbf{U}^T : (\nu : \mathbf{Sub} \Omega \Gamma) \rightarrow \mathbf{Tm} \Omega \mathbf{U} \rightarrow \mathbf{Set}_{j+1}$$

$$\mathbf{U}^T \nu a \equiv \mathbf{Tm} \Omega (\mathbf{El} a)$$

$$(\mathbf{El} a)^T : (\nu : \mathbf{Sub} \Omega \Gamma)(t : \mathbf{Tm} \Omega (\mathbf{El} (a[\nu]))) \rightarrow a^A (\Gamma^T \nu)$$

$$(\mathbf{El} a)^T \nu t \equiv (a^T \nu) t$$

In the interpretation of **El**, note that

$$a^T \nu : \mathbf{Tm} \Omega (\mathbf{El} (a[\nu])) \simeq_{\text{id}} a^A (\Gamma \nu)$$

But this is an isomorphism in the semantic \mathbf{U} , which is the category of sets in \mathbf{Set}_{j+1} . So coercion along $a^T \nu$ is simply function application, and we are justified in writing $(a^T \nu) t$.

For each type former in \mathbf{U} , we have to exhibit an isomorphism of sets in the interpretation.

\top, Σ

We need

$$\top^T : (\nu : \mathbf{Sub} \Omega \Gamma) \rightarrow \mathbf{U}^T \nu (\top[\nu]) \simeq_{\text{id}} \top^A (\Gamma^T \nu)$$

The result type computes to $\mathbf{Tm} \Omega (\mathbf{El} \top) \simeq \top$, which is evident. For Σ , we have to show

$$\mathbf{Tm} \Omega (\mathbf{El} (\Sigma (a[\nu]) (b[\nu \circ \mathbf{p}, \mathbf{q}]))) \simeq ((\alpha : a^A (\Gamma^T \nu)) \times b^A (\Gamma^T \nu, \alpha))$$

This follows from the induction hypotheses a^T and b^T , which establish the first and second components of the desired isomorphism.

Identity

For the identity type, we need

$$\mathbf{Tm} \Omega (\mathbf{El} (\mathbf{Id} (t[\nu]) (u[\nu]))) \simeq (t^A (\Gamma^T \nu) \equiv u^A (\Gamma^T \nu))$$

This follows from $t^T \nu$, $u^T \nu$ and the specifying isomorphism of \mathbf{Id} .

Infinitary function

This function type follows the same pattern. We define the isomorphism below using induction hypotheses and the specifying isomorphism of Π^{inf} .

$$\mathbf{Tm} \Omega (\mathbf{El} (\Pi^{\text{inf}} Ix (\lambda i. (b i)[\nu]))) \simeq ((i : Ix) \rightarrow (b i)^A (\Gamma \nu))$$

Inductive function

Inductive functions are interpreted using transport along $a^T \nu : \mathbf{Tm} \Omega (\mathbf{El} (a[\nu])) \simeq a^A (\Gamma^T \nu)$:

$$\begin{aligned} (\Pi a B)^T : (\nu : \mathbf{Sub} \Omega \Gamma) & (t : \mathbf{Tm} \Omega (\Pi (a[\nu]) (B[\nu \circ \mathbf{p}, \mathbf{q}]))) \\ & \rightarrow (\alpha : a^A (\Gamma^T \nu)) \rightarrow B^A (\Gamma^T \nu, \alpha) \\ (\Pi a B)^T \nu t & \equiv \lambda \alpha. B (\nu, (a^T \nu)^{-1} \alpha) (t ((a^T \nu)^{-1} \alpha)) \end{aligned}$$

External functions are interpreted the same way as in the finitary case.

5.4.2 Eliminator Construction

We only present the eliminator construction in the following, since (unique) recursors are derivable from this.

Compared to the finitary case, the eliminator construction does not change as much as the term algebra construction. The reason is that although we have weakened strict algebra equality to isomorphism, in the current construction we only have to show equalities of substitutions and terms, which we do not need to weaken (and they cannot be sensibly weakened anyway).

We assume j and k such that $j + 1 \leq k$, and also $\Omega : \mathbf{Sig}_j$ and $\omega^D : \Omega_k^D(\Omega^T \text{id})$. Hence, ω^D is a displayed Ω -algebra over the term algebra, and we aim to construct its section. Note that we lift $\Omega^T \text{id} : \Omega_{j+1}^A$ to level k by cumulativity. We define $-^E$ by induction on \mathbf{syn}_j .

$$\begin{aligned} -^E : (\Gamma : \mathbf{Con}) \quad (\nu : \mathbf{Sub} \Omega \Gamma) &\rightarrow \Gamma^S(\nu^A(\Omega^T \text{id}))(\nu^D \omega^D) \\ -^E : (\sigma : \mathbf{Sub} \Gamma \Delta)(\nu : \mathbf{Sub} \Omega \Gamma) &\rightarrow \Delta^E(\sigma \circ \nu) \equiv \sigma^S(\Gamma^E \nu) \\ -^E : (A : \mathbf{Ty} \Gamma) \quad (\nu : \mathbf{Sub} \Omega \Gamma)(t : \mathbf{Tm} \Omega(A[\nu])) &\rightarrow A^S(t^A(\Omega^T \text{id}))(t^D \omega^D)(\Gamma^E \nu) \\ -^E : (t : \mathbf{Tm} \Gamma A) \quad (\nu : \mathbf{Sub} \Omega \Gamma) &\rightarrow A^E \nu(t[\nu]) \equiv t^S(\Gamma^E \nu) \end{aligned}$$

This is so far exactly the same as in Section 4.4.5. The subsequent changes arise from the need to transport along $-^T$ in definitions.

Universe

For the universe, we need

$$\mathbf{U}^E : (\nu : \mathbf{Sub} \Omega \Gamma)(a : \mathbf{Tm} \Omega \mathbf{U}) \rightarrow (\alpha : a^A(\Omega^T \text{id})) \rightarrow a^D \omega^D \alpha$$

Since we only have $a^T \text{id} : a^A(\Omega^T \text{id}) \simeq \mathbf{Tm} \Omega(\text{El } a)$, the definition becomes

$$\mathbf{U}^E \nu a t \equiv (a^T \text{id } t)^D \omega^D$$

That this is well-typed, follows from

$$\begin{aligned} ((a^T \text{id}) t)^T \text{id} &: t \equiv ((a^T \text{id}) t)^A(\Omega^T \text{id}) \\ ((a^T \text{id}) t)^D \omega^D &: a^D \omega^D (((a^T \text{id}) t)^A(\Omega^T \text{id})) \end{aligned}$$

For El , we need to show

$$(\text{El } a)^E : (\nu : \mathbf{Sub} \Omega \Gamma)(t : \mathbf{Tm} \Omega(\text{El}(a[\nu]))) \rightarrow a^S(\Gamma^E \nu)(t^A(\Omega^T \text{id})) \equiv t^D \omega^D$$

We have

$$t^T \text{id} : (a[\nu])^T \text{id} t \equiv t^A (\Omega^T \text{id})$$

Moreover

$$a^E \nu : \mathbf{U}^E \nu (a[\nu]) \equiv a^S (\Gamma^E \nu)$$

Hence

$$a^E \nu : (\lambda t. ((a[\nu])^T \text{id} t)^D \omega^D) \equiv a^S (\Gamma^E \nu)$$

Applying both sides to $((a[\nu])^T \text{id})^{-1} t$, we have

$$((a[\nu])^T \text{id} (((a[\nu])^T \text{id})^{-1} t))^D \omega^D \equiv a^S (\Gamma^E \nu) (((a[\nu])^T \text{id})^{-1} t)$$

This simplifies to

$$t^D \omega^D \equiv a^S (\Gamma^E \nu) (((a[\nu])^T \text{id})^{-1} t)$$

By $(a^T \text{id} t)^T \text{id} : t \equiv (a^T \text{id} t)^A (\Omega^T \text{id})$ this becomes:

$$t^D \omega^D \equiv a^S (\Gamma^E \nu) (((a[\nu])^T \text{id})^{-1} ((a^T \text{id} t)^A (\Omega^T \text{id})))$$

Thus we have the required

$$t^D \omega^D \equiv a^S (\Gamma^E \nu) (\Omega^T \text{id})$$

\top, Σ

For \top , we need

$$\top^E : (\nu : \mathbf{Sub} \Omega \Gamma) \rightarrow \mathbf{U}^E \nu \top \equiv \top^S (\Gamma^E \nu)$$

But this is clearly trivial, since $\top^S (\Gamma^E \nu) : \top \rightarrow \top$. Considering Σ :

$$(\Sigma a b)^E : (\nu : \mathbf{Sub} \Omega \Gamma) \rightarrow \mathbf{U}^E \nu (\Sigma (a[\nu]) (b[\nu \circ \mathbf{p}, \mathbf{q}])) \equiv (\Sigma a b)^S (\Gamma^E \nu)$$

This case is a bit tedious. The sides above are functions, so appealing to function extensionality we apply both sides to (α, β) , where $\alpha : a^A (\nu^A (\Omega^T \text{id}))$ and $\beta : b^A (\nu^A (\Omega^T \text{id}), \alpha)$. We also unfold some definitions:

$$(((\Sigma (a[\nu]) (b[\nu \circ \mathbf{p}, \mathbf{q}]))^T \text{id})^{-1} (\alpha, \beta))^D \omega^D \equiv (a^S (\Gamma^E \nu) \alpha, b^S (\Gamma^E \nu, \text{refl}) \beta)$$

Unfolding the left side of this equation, we have

$$((((a[\nu])^T \text{id})^{-1} \alpha)^D \omega^D, (((b[\nu \circ \mathbf{p}, \mathbf{q}])^T (\text{id}, ((a[\nu])^T \text{id})^{-1} \alpha))^{-1} \beta)^D \omega^D)$$

Let us abbreviate $((a[\nu])^T \text{id})^{-1} \alpha : \mathbf{Tm} \Omega (\text{El} (a[\nu]))$ as α' :

$$(\alpha'^D \omega^D, (((b[\nu \circ \mathbf{p}, \mathbf{q}])^T (\text{id}, \alpha'))^{-1} \beta)^D \omega^D)$$

Hence, we need to show component-wise equality of pairs. The equality of first components follow from the following:

$$a^E \nu : \mathbf{U}^E \nu (a[\nu]) \equiv a^S (\Gamma^E \nu)$$

Unfolding definitions and applying both sides to α , we the equality of first components:

$$\alpha'^D \omega^D \equiv a^S (\Gamma^E \nu) \alpha$$

Analogously, the equality of second components follows from

$$b^E (\nu, \alpha') : (((b[\nu, \alpha'])^T \text{id})^{-1} \beta)^D \omega^D \equiv b^S (\Gamma^E \nu, \text{refl}) \beta$$

The right hand side is what we need, the left hand side though does not immediately match up. Hence, it remains to show that

$$(((b[\nu, \alpha'])^T \text{id})^{-1} \beta)^D \omega^D \equiv (((b[\nu \circ \mathbf{p}, \mathbf{q}])^T (\text{id}, \alpha'))^{-1} \beta)^D \omega^D$$

Thus, it suffices to show

$$(b[\nu, \alpha'])^T \text{id} \equiv (b[\nu \circ \mathbf{p}, \mathbf{q}])^T (\text{id}, \alpha')$$

This equation follows from a somewhat laborious unfolding of all involved definitions. In particular, we use that for some $a : \mathbf{Tm} \Gamma \mathbf{U}$, we have

$$(a[\sigma])^T \nu \equiv a^M (\sigma^T \nu) \circ a^T (\sigma \circ \nu)$$

which follows from the definition of $-^T$.

Inductive functions

In Π we likewise transport along the domain isomorphism.

$$\begin{aligned} (\Pi a B)^E : (\nu : \mathbf{Sub} \Omega \Gamma)(t : \mathbf{Tm} \Omega (\Pi (a[\nu]) (B[\nu \circ \mathbf{p}, \mathbf{q}]))) \\ \rightarrow (\alpha : \mathbf{Tm} \Omega (\text{El} (a[\nu]))) \rightarrow B^S (t^A (\Omega^T \text{id}) \alpha) (t^D \omega^D (\alpha^D \omega^D)) (\Gamma^E \nu, \text{refl}) \\ (\Pi a B)^E \nu t \equiv \lambda \alpha. B^E (\nu, (a[\nu])^T \text{id} \alpha) (t ((a[\nu])^T \text{id} \alpha)) \end{aligned}$$

This is well-typed by the following:

$$\begin{aligned} a^E \nu & : ((a[\nu])^T \text{id} \alpha)^D \omega^D \equiv a^S (\Gamma^E \nu) \alpha \\ ((a[\nu])^T \text{id} \alpha)^T \text{id} : \alpha & \equiv ((a[\nu])^T \text{id} \alpha)^A (\Omega^T \text{id}) \end{aligned}$$

Id, Π^{inf} , Π^{ext}

Id is trivial by UIP, and for Π^{inf} and Π^{ext} we again do a straightforward recursion under the indexing function.

This concludes the definition of $-^E$. We again show the initiality of term algebras.

Definition 62 (Eliminators). Assuming $\Omega : \text{Sig}_j$, a k level such that $k \geq j + 1$ and $\omega^D : \Omega_k^D (\Omega^T \text{id})$, we have $\Omega^E \text{id} : \Omega^S (\Omega^T \text{id}) \omega^D$ as the eliminator.

Theorem 8. $\Omega^T \text{id} : \Omega_{j+1}^A$ is initial when lifted to any $k \geq j + 1$ level.

Proof. $\Omega^T \text{id}$ supports elimination by Definition 62, and elimination is equivalent to initiality by Theorem 1. \square

5.5 Levitation and Bootstrapping

In this section we adapt the bootstrapping procedure from Section 4.5 to infinitary signatures.

Bootstrapping for 2LTT semantics

If we only want to write down signatures and get their 2LTT-based semantics, a simplified bootstrapping suffices, which is essentially the same as in Section 4.5. We write $\text{ToS}_i : \text{Set}_{i+1}$ for the type of models where underlying sets are in Set_i and external indexing is over Ty_0 . We also have $\mathbf{M}_i : \text{ToS}_{i+2}$ for the flwf models where underlying sets in algebras are in Set_i and external indexing is over types in Ty_0 .

Definition 63. The type of **bootstrap signatures** is defined as follows:

$$\text{BootSig} \equiv (i : \text{Level}) \rightarrow (M : \text{ToS}_i) \rightarrow \text{Con}_M$$

These bootstrap signatures only allow external indexing by types in Ty_0 . We can write bootstrap signatures and interpret them in \mathbf{M}_i , by applying them to \mathbf{M}_i .

Bootstrapping for term algebras

Now we reuse the ETT setting from Section 4.4.2. We have $\text{ToS}_{i,j} : \text{Set}_{i+1 \sqcup j+1}$ for the type of models where underlying sets are in Set_i and Π^{inf} and Π^{ext} abstract over Set_j . We also have $\mathbf{M}_{i,j} : \text{ToS}_{(i+1 \sqcup j)+1,j}$ as the flwf models, again with underlying sets of algebras in Set_i and external indexing types in Set_j .

Definition 64. The type of **bootstrap signatures** at level j is defined as follows. These may contain external indexing by types in Set_j .

$$\text{BootSig}_j := (i, j : \text{Level}) \rightarrow (M : \text{ToS}_{i,j}) \rightarrow \text{Con}_M$$

Definition 65 (Signature for ToS). We define $\text{ToSSig}_j : \text{BootSig}_{j+1}$ as the bootstrap signature for ToS, where the described signatures may be indexed by types in Set_j . Like in Section 4.5, we use an internal notation. We present an excerpt.

```

Con  : U
Sub  : Con → Con → U
Ty   : Con → U
Tm   : (Γ : Con) → Ty Γ → U
...
SigU : {Γ : Con} → El (Ty Γ)
SigEl : {Γ : Con} → Tm Γ SigU → El (Ty Γ)
Πinf : {Γ : Con} (A : Setj) →ext (A →inf Tm Γ SigU) → El (Tm Γ SigU)
...

```

Now the interpretation of ToSSig_j in $\mathbf{M}_{i,j+1}$ yields the flwf where objects are elements of $\text{ToS}_{i,j}$. Note the level bump: ToSSig_j is in BootSig_{j+1} , so we expend one level at each round of self-description. We get the notion of ToS-induction from $\mathbf{M}_{i,j+1}$, and we have $\text{ToS}_{i,j} \leq \text{ToS}_{i+1,j}$ (by definition of ToS and the rules of subtyping), which allows us to specify what it means for a model to support elimination into any universe. Thus we recover all concepts that are used in the term algebra and eliminator constructions.

5.6 Left Adjoints of Substitutions

TODO

CHAPTER 6

Higher Inductive-Inductive Signatures

So far we only considered semantics of signatures where equality constructors are interpreted as proof-irrelevant equalities, i.e. those satisfying UIP. This inspires the naming of *quotient* inductive-inductive signatures. In contrast, *higher inductive-inductive* signatures are characterized by having possibly proof-relevant and iterated equalities in algebras. The natural setting of HIITs is homotopy type theory (HoTT) [Uni13], where higher equalities are available, and unlike in basic intensional type theory, they can be also used in interesting ways, thanks to univalence.

In short, we might think of HIITs as QIITs in homotopy type theory, or alternatively, QIITs are specific HIITs that are set-truncated. Thus HIITs are more general than QIITs, modulo some minor details.

We shall see that 2LTT is firmly necessary in the specification of HIITs, and the ToS must live in the outer layer. The reason is that there is no known way to sensibly internalize the metatheory of type theories purely inside HoTT. This is the problem of “HoTT eating itself” [Shu14]. It is also closely related to the problem of representing semisimplicial types purely inside HoTT. If we can construct semisimplicial types in an embedded type theory, and interpret that into non-truncated HoTT types, that would indeed solve the problem. But so far it has not been solved, or proven impossible. A main point of 2LTT is precisely to allow construction of semisimplicial types [ACKS19].

Overview

This chapter is based on the author’s joint work with Kaposi [KK20]. In the following, we specify the ToS and give a semantics of signatures which is more

limited than what we had before. Concretely, we restrict ourselves to the AMDS fragment of semantics. This is sufficient to compute what we mean by induction and initiality (which has been called “homotopy initiality” in the context of HoTT [Soj15]).

Why don’t we go further? The main reason is that the natural semantics is actually in $(\omega, 1)$ -categories: we want $(\omega, 1)$ -categories of algebras. This requires a different approach and toolset. In particular, in [KK20] we gave an example that a naive attempt to extend the AMDS semantics of signatures with the notion of identity morphisms already fails. The author of this thesis is not versed enough in higher category theory, so we leave the exposition of the full semantics to future work.

We do note that the higher semantics has been done by Capriotti and Sattler. See [CS] for an abstract; the bulk of the work remains unpublished as of now. In short, Capriotti and Sattler define ToS in 2LTT, and also use 2LTT to give a model where signatures are higher categories specified as complete Segal types. They show that categories of algebras have finite limits and that initiality is equivalent to induction. Additionally, the setup yields a structure identity principle for each signature. However, reductions to simpler type formers are not discussed, nor possible term algebra constructions. Both of these appear to be far more difficult than in the quotient setting, and to the author’s knowledge there are no concrete proposals how to approach them.

Metatheory

We work in 2LTT. We assume that Ty_0 is closed under Π , Σ , \top and *intensional* identity $- = -$. Thus, unlike before, the inner theory is intensional and compatible with univalence. We use the “based” induction principle for $- = -$ [Uni13, Section 1.12.1]:

$$\begin{aligned} \mathsf{J} : \{A : \mathsf{Ty}_0\} \{x : A\} (B : (y : A) \rightarrow x = y \rightarrow \mathsf{Ty}_0) &\rightarrow B \ x \ \mathsf{refl} \\ &\rightarrow (y : A) (p : x = y) \rightarrow B \ y \ p \\ \mathsf{J}\beta : \mathsf{J} \ B \ br \ x \ \mathsf{refl} &\equiv br \end{aligned}$$

6.1 Theory of Signatures

The theory of HII signatures is fairly similar to the theory of infinitary QII signatures. The primary difference is that **ld** is intensional in the ToS too: it has **refl** and **J**, and does not have equality reflection.

However, the choice of β -rules is now a bit more delicate. We have two options in the specification of the ToS.

1. One choice is to have definitional β -rules for **ld** and Π , expressed using $- \equiv -$. However, this forces us to have strict equalities in the $-^M$ and $-^S$ interpretations of **El**, in other words, to have algebra morphisms and sections which preserve structure definitionally. As we will see, if we instead use inner (“fibrant”) equality for structure-preservation, the $-^M$ and $-^S$ interpretations fail to respect the mentioned β -rules in the ToS.

Now, such strict morphisms and sections of algebras are sensible and useful when we want to specify HIITs, because β -rules of recursors and eliminators are generally expected to be definitional in HoTT. But they are not appropriate when we want to internally talk about algebra morphisms in HoTT, as strict equations are only available in 2LTT. For example, it only makes sense to specify a structure identity principle with “weak” algebra morphisms, which preserve structure up to inner equality.

2. The other option is to drop strict β for Π and **ld**. In this case, we can still specify β for **ld** using **ld** itself, but we cannot specify β and η for Π at all, because **ld** can only express equality of small terms in some $\mathsf{Tm} \Gamma (\mathsf{El} a)$. We *could* add a “big” version of **ld** for this purpose, but this feels like an unnecessary complication. Instead, for the sake of simplicity, we just drop **lam** for Π together with the $\beta\eta$ rules. This has a negligible impact on the practical expressiveness of signatures, although the ToS expressiveness for specifying model constructions does take a hit.

In summary, we only have **app** for Π , and a weak β -rule for **J**. This works fine in the semantics, as $-^M$ and $-^S$ do preserve weak $\mathsf{J}\beta$ up to inner equality.

We shall refer to option 1 as **strict ToS**, and option 2 as **weak ToS**. In the following we present the two options simultaneously.

Definition 66. A **model of ToS** (weak or strict) consists of a base cwf, with underlying sets returning in **Set**, supporting the following type formers.

- A **cwf** with underlying sets **Con**, **Sub**, **Ty** and **Tm**, all returning in the outer **Set** universe of 2LTT.
- A **Tarski-style universe** \mathbf{U} with decoding \mathbf{El} . \mathbf{U} is closed under the following type formers:
 - The **unit type** \top .
 - Σ -**types** $\Sigma : (a : \mathbf{Tm} \Gamma \mathbf{U}) \rightarrow \mathbf{Tm} (\Gamma \triangleright \mathbf{El} a) \mathbf{U} \rightarrow \mathbf{Tm} \Gamma \mathbf{U}$, with specifying isomorphism

$$(\text{proj}, -, -) : \mathbf{Tm} \Gamma (\mathbf{El} (\Sigma a b)) \simeq (t : \mathbf{Tm} \Gamma (\mathbf{El} a)) \times \mathbf{Tm} \Gamma (\mathbf{El} (b[\text{id}, t]))$$
 - **Extensional identity** $\text{Id} : \mathbf{Tm} \Gamma (\mathbf{El} a) \rightarrow \mathbf{Tm} \Gamma (\mathbf{El} a) \rightarrow \mathbf{Tm} \Gamma \mathbf{U}$, specified by $(\text{reflect}, \text{refl}) : \mathbf{Tm} \Gamma (\mathbf{El} (\text{Id } t u)) \simeq (t \equiv u)$.
 - **Infinitary functions** $\Pi^{\text{inf}} : (A : \mathbf{Ty}_0) \rightarrow (A \rightarrow \mathbf{Tm} \Gamma \mathbf{U}) \rightarrow \mathbf{Tm} \Gamma \mathbf{U}$, specified by $(\text{app}^{\text{inf}}, \text{lam}^{\text{inf}}) : \mathbf{Tm} \Gamma (\Pi^{\text{inf}} A b) \simeq ((\alpha : A) \rightarrow \mathbf{Tm} \Gamma (\mathbf{El} (b \alpha)))$.
- An **inductive function type** $\Pi : (a : \mathbf{Tm} \Gamma \mathbf{U}) \rightarrow \mathbf{Tm} (\Gamma \triangleright \mathbf{El} a) \rightarrow \mathbf{Tm} \Gamma$, specified by $(\text{app}, \text{lam}) : \mathbf{Tm} \Gamma (\Pi a B) \simeq \mathbf{Tm} (\Gamma \triangleright \mathbf{El} a) B$.
- An **external function type** $\Pi^{\text{ext}} : (A : \mathbf{Ty}_0) \rightarrow (A \rightarrow \mathbf{Tm} \Gamma) \rightarrow \mathbf{Tm} \Gamma$, specified by $(\text{app}^{\text{ext}}, \text{lam}^{\text{ext}}) : \mathbf{Tm} \Gamma (\Pi^{\text{ext}} A B) \simeq ((\alpha : A) \rightarrow \mathbf{Tm} \Gamma (B \alpha))$.

6.2 Semantics

CHAPTER 7

Conclusion

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